**An Introduction to Distributed Algorithms**  
by Valmir C. Barbosa  
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A senior undergraduate or graduate level computer science textbook on algorithm design for distributed computer systems.

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**Synopsis**  
by Dean Andrews

Distributed computing poses special challenges for software developers. When programs must run across multiple processors, either multiprocessors within the same computer or processors distributed across a computer network, software developers encounter the unique problems of information propagation, synchronization, deadlock detection, and more. In *An Introduction to Distributed Algorithms*, author Valmir Barbosa describes general use algorithms (for any language or platform) to meet the demands of distributed software design. He also devotes chapters to topics like program debugging and simulation that are seldom covered in other books. Each of the ten chapters ends with a bibliography and exercises.

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An Introduction to Distributed Algorithms takes up some of the main concepts and algorithms, ranging from basic to advanced techniques and applications, that underlie the programming of distributed-memory systems such as computer networks, networks of workstations, and multiprocessors. Written from the broad perspective of distributed-memory systems in general, it includes topics such as algorithms for maximum flow, program debugging, and simulation that do not appear in other texts on distributed algorithms.

Moving from fundamentals to advances and applications, ten chapters -- with exercises and bibliographic notes -- cover a variety of topics. These include models of distributed computation, information propagation, leader election, distributed snap-shots, network synchronization, self-stability, termination detection, deadlock detection, graphic algorithms, mutual exclusion, program debugging, and simulation.

All of the algorithms are present in a clear, template-based format for the description of message-passing computations among the nodes of a connected graph. Such a generic setting allows the treatment of problems originating from many different application areas. The main ideas and algorithms are described in a way that balances intuition and formal rigor -- most are preceded by general intuitive discussion and followed by formal statements as to correctness, complexity, or other properties.

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Preface

This book presents an introduction to some of the main problems, techniques, and algorithms underlying the programming of distributed-memory systems, such as computer networks, networks of workstations, and multiprocessors. It is intended mainly as a textbook for advanced undergraduates or first-year graduate students in computer science and requires no specific background beyond some familiarity with basic graph theory, although prior exposure to the main issues in concurrent programming and computer networks may
also be helpful. In addition, researchers and practitioners working on distributed computing will also find it useful as a general reference on some of the most important issues in the field.

The material is organized into ten chapters covering a variety of topics, such as models of distributed computation, information propagation, leader election, distributed snapshots, network synchronization, self-stability, termination detection, deadlock detection, graph algorithms, mutual exclusion, program debugging, and simulation. Because I have chosen to write the book from the broader perspective of distributed-memory systems in general, the topics that I treat fail to coincide exactly with those normally taught in a more orthodox course on distributed algorithms. What this amounts to is that I have included topics that normally would not be touched (as algorithms for maximum flow, program debugging, and simulation) and, on the other hand, have left some topics out (as agreement in the presence of faults).

All the algorithms that I discuss in the book are given for a "target" system that is represented by a connected graph, whose nodes are message-driven entities and whose edges indicate the possibilities of point-to-point communication. This allows the algorithms to be presented in a very simple format by specifying, for each node, the actions to be taken to initiate participating in the algorithm and upon the receipt of a message from one of the nodes connected to it in the graph. In describing the main ideas and algorithms, I have sought a balance between intuition and formal rigor, so that most are preceded by a general intuitive discussion and followed by formal statements regarding correctness, complexity, or other properties.

The book's ten chapters are grouped into two parts. Part 1 is devoted to the basics in the field of distributed algorithms, while Part 2 contains more advanced techniques or applications that build on top of techniques discussed previously. Part 1 comprises Chapters 1 through 5. Chapters 1 and 2 are introductory chapters, although in two different ways. While Chapter 1 contains a discussion of various issues related to message-passing systems that in the end lead to the adoption of the generic message-driven system I mentioned earlier, Chapter 2 is devoted to a discussion of constraints that are inherent to distributed-memory systems, chiefly those related to a system's asynchronism or synchronism, and the anonymity of its constituents. The remaining three chapters of Part 1 are each dedicated to a group of fundamental ideas and techniques, as follows. Chapter 3 contains models of computation and complexity measures, while Chapter 4 contains some fundamental algorithms (for information propagation and some simple graph problems) and Chapter 5 is devoted to fundamental techniques (as leader election, distributed snapshots, and network synchronization).

The chapters that constitute Part 2 are Chapters 6 through 10. Chapter 6 brings forth the subject of stable properties, both from the perspective of selfstability and of stability detection (for termination and deadlock detection). Chapter 7 contains graph algorithms for minimum spanning trees and maximum flows. Chapter 8 contains algorithms for resource sharing under the requirement of mutual exclusion in a variety of circumstances, including generalizations of the paradigmatic dining philosophers problem. Chapters 9 and 10 are, respectively, dedicated to the topics of program debugging and simulation. Chapter 9 includes techniques for program re-execution and for breakpoint detection. Chapter 10 deals with time-stepped simulation, conservative event-driven simulation, and optimistic event-driven simulation.

Every chapter is complemented by a section with exercises for the reader and another with bibliographic notes. Of the exercises, many are intended to bring the reader one step further in the treatment of some topic discussed in the chapter. When this is the case, an indication is given, during the discussion of the topic, of the exercise that may be pursued to expand the treatment of that particular topic. I have attempted to collect a fairly comprehensive set of bibliographic references, and the sections with bibliographic notes are intended to provide the reader with the source references for the main issues treated in the chapters, as well as to indicate how to proceed further.

I believe the book is sized reasonably for a one-term course on distributed algorithms. Shorter syllabi are also possible, though, for example by omitting Chapters 1 and 2 (except
for Sections 1.4 and 2.1), then covering Chapters 3 through 6 completely, and then selecting as many chapters as one sees fit from Chapters 7 through 10 (the only interdependence that exists among these chapters is of Section 10.2 upon some of Section 8.3).

Notation
The notation $\log^k n$ is used to indicate $(\log n)^k$. All of the remaining notation in the book is standard.

Part 1: Fundamentals

Message-Passing Systems
Intrinsic Constraints
Models of Computation
Basic Algorithms
Basic Techniques

Part Overview

This first part of the book is dedicated to some of the fundamentals in the field of distributed algorithms. It comprises five chapters, in which motivation, some limitations, models, basic algorithms, and basic techniques are discussed.

Chapter 1 opens with a discussion of the distributed-memory systems that provide the motivation for the study of distributed algorithms. These include computer networks, networks of workstations, and multiprocessors. In this context, we discuss some of the issues that relate to the study of those systems, such as routing and flow control, message buffering, and processor allocation. The chapter also contains the description of a generic template to write distributed algorithms, to be used throughout the book.

Chapter 2 begins with a discussion of full asynchronism and full synchronism in the context of distributed algorithms. This discussion includes the introduction of the asynchronous and synchronous models of distributed computation to be used in the remainder of the book, and the presentation of details on how the template introduced in Chapter 1 unfolds in each of the two models. We then turn to a discussion of intrinsic limitations in the context of anonymous systems, followed by a brief discussion of the notions of knowledge in distributed computations.

The computation models introduced in Chapter 2 (especially the asynchronous model) are in Chapter 3 expanded to provide a detailed view in terms of events, orders, and global states. This view is necessary for the proper treatment of timing issues in distributed computations, and also allows the introduction of the complexity measures to be employed throughout. The chapter closes with a first discussion (to be resumed later in Chapter 5) of how the asynchronous and synchronous models relate to each other.

Chapters 4 and 5 open the systematic presentation of distributed algorithms, and of their properties, that constitutes the remainder of the book. Both chapters are devoted to basic material. Chapter 4, in particular, contains basic algorithms in the context of information propagation and of some simple graph problems.

In Chapter 5, three fundamental techniques for the development of distributed algorithms are introduced. These are the techniques of leader election (presented only for some types of systems, as the topic is considered again in Part 2, Chapter 7), distributed snapshots, and network synchronization. The latter two techniques draw heavily on material introduced earlier in Chapter 3, and constitute some of the essential building blocks to be occasionally used in later chapters.

Chapter 1: Message-Passing Systems

Overview
The purpose of this chapter is twofold. First we intend to provide an overall picture of various real-world sources of motivation to study message-passing systems, and in doing so to provide the reader with a feeling for the several characteristics that most of those systems
share. This is the topic of Section 1.1, in which we seek to bring under a same framework seemingly disparate systems as multiprocessors, networks of workstations, and computer networks in the broader sense.

Our second main purpose in this chapter is to provide the reader with a fairly rigorous, if not always realizable, methodology to approach the development of message-passing programs. Providing this methodology is a means of demonstrating that the characteristics of real-world computing systems and the main assumptions of the abstract model we will use throughout the remainder of the book can be reconciled. This model, to be described timely, is graph-theoretic in nature and encompasses such apparently unrealistic assumptions as the existence of infinitely many buffers to hold the messages that flow on the system's communication channels (thence the reason why reconciling the two extremes must at all be considered).

This methodology is presented as a collection of interrelated aspects in Sections 1.2 through 1.7. It can also be viewed as a means to abstract our thinking about message-passing systems from various of the peculiarities of such systems in the real world by concentrating on the few aspects that they all share and which constitute the source of the core difficulties in the design and analysis of distributed algorithms. Sections 1.2 and 1.3 are mutually complementary, and address respectively the topics of communication processors and of routing and flow control in message-passing systems. Section 1.4 is devoted to the presentation of a template to be used for the development of message-passing programs. Among other things, it is here that the assumption of infinite-capacity channels appears. Handling such an assumption in realistic situations is the topic of Section 1.5. Section 1.6 contains a treatment of various aspects surrounding the question of processor allocation, and completes the chapter's presentation of methodological issues. Some remarks on some of the material presented in previous sections comes in Section 1.7. Exercises and bibliographic notes follow respectively in Sections 1.8 and 1.9.

1.1 Distributed-memory systems

Message passing and distributed memory are two concepts intimately related to each other. In this section, our aim is to go on a brief tour of various distributed-memory systems and to demonstrate that in such systems message passing plays a chief role at various levels of abstraction, necessarily at the processor level but often at higher levels as well.

Distributed-memory systems comprise a collection of processors interconnected in some fashion by a network of communication links. Depending on the system one is considering, such a network may consist of point-to-point connections, in which case each communication link handles the communication traffic between two processors exclusively, or it may comprise broadcast channels that accommodate the traffic among the processors in a larger cluster. Processors do not physically share any memory, and then the exchange of information among them must necessarily be accomplished by message passing over the network of communication links.

The other relevant abstraction level in this overall panorama is the level of the programs that run on the distributed-memory systems. One such program can be thought of as comprising a collection of sequential-code entities, each running on a processor, maybe more than one per processor. Depending on peculiarities well beyond the intended scope of this book, such entities have been called tasks, processes, or threads, to name some of the denominations they have received. Because the latter two forms often acquire context-dependent meanings (e.g., within a specific operating system or a specific programming language), in this book we choose to refer to each of those entities as a task, although this denomination too may at times have controversial connotations.

While at the processor level in a distributed-memory system there is no choice but to rely on message passing for communication, at the task level there are plenty of options. For example, tasks that run on the same processor may communicate with each other either through the explicit use of that processor's memory or by means of message passing in a very natural way. Tasks that run on different processors also have essentially these two
possibilities. They may communicate by message passing by relying on the message-passing mechanisms that provide interprocessor communication, or they may employ those mechanisms to emulate the sharing of memory across processor boundaries. In addition, a myriad of hybrid approaches can be devised, including for example the use of memory for communication by tasks that run on the same processor and the use of message passing among tasks that do not.

Some of the earliest distributed-memory systems to be realized in practice were long-haul computer networks, i.e., networks interconnecting processors geographically separated by considerable distances. Although originally employed for remote terminal access and somewhat later for electronic-mail purposes, such networks progressively grew to encompass an immense variety of data-communication services, including facilities for remote file transfer and for maintaining work sessions on remote processors. A complex hierarchy of protocols is used to provide this variety of services, employing at its various levels message passing on point-to-point connections. Recent advances in the technology of these protocols are rapidly leading to fundamental improvements that promise to allow the coexistence of several different types of traffic in addition to data, as for example voice, image, and video. The protocols underlying these advances are generally known as Asynchronous Transfer Mode (ATM) protocols, in a way underlining the aim of providing satisfactory service for various different traffic demands. ATM connections, although frequently of the point-to-point type, can for many applications benefit from efficient broadcast capabilities, as for example in the case of teleconferencing.

Another notorious example of distributed-memory systems comes from the field of parallel processing, in which an ensemble of interconnected processors (a multiprocessor) is employed in the solution of a single problem. Application areas in need of such computational potential are rather abundant, and come from various of the scientific and engineering fields. The early approaches to the construction of parallel processing systems concentrated on the design of shared-memory systems, that is, systems in which the processors share all the memory banks as well as the entire address space. Although this approach had some success for a limited number of processors, clearly it could not support any significant growth in that number, because the physical mechanisms used to provide the sharing of memory cells would soon saturate during the attempt at scaling.

The interest in providing massive parallelism for some applications (i.e., the parallelism of very large, and scalable, numbers of processors) quickly led to the introduction of distributed-memory systems built with point-to-point interprocessor connections. These systems have dominated the scene completely ever since. Multiprocessors of this type were for many years used with a great variety of programming languages endowed with the capability of performing message passing as explicitly directed by the programmer. One problem with this approach to parallel programming is that in many application areas it appears to be more natural to provide a unique address space to the programmer, so that, in essence, the parallelization of preexisting sequential programs can be carried out in a more straightforward fashion. With this aim, distributed-memory multiprocessors have recently appeared whose message-passing hardware is capable of providing the task level with a single address space, so that at this level message passing can be done away with. The message-passing character of the hardware is fundamental, though, as it seems that this is one of the key issues in providing good scalability properties along with a shared-memory programming model. To provide this programming model on top of a message-passing hardware, such multiprocessors have relied on sophisticated cache techniques.

The latest trend in multiprocessor design emerged from a re-consideration of the importance of message passing at the task level, which appears to provide the most natural programming model in various situations. Current multiprocessor designers are then attempting to build, on top of the message-passing hardware, facilities for both message-passing and scalable shared-memory programming.
As our last example of important classes of distributed-memory systems, we comment on networks of workstations. These networks share a lot of characteristics with the long-haul networks we discussed earlier, but unlike those they tend to be concentrated within a much narrower geographic region, and so frequently employ broadcast connections as their chief medium for interprocessor communication (point-to-point connections dominate at the task level, though). Also because of the circumstances that come from the more limited geographic dispersal, networks of workstations are capable of supporting many services other than those already available in the long-haul case, as for example the sharing of file systems. In fact, networks of workstations provide unprecedented computational and storage power in the form, respectively, of idling processors and unused storage capacity, and because of the facilitated sharing of resources that they provide they are already beginning to be looked at as a potential source of inexpensive, massive parallelism.

As it appears from the examples we described in the three classes of distributed-memory systems we have been discussing (computer networks, multiprocessors, and networks of workstations), message-passing computations over point-to-point connections constitute some sort of a pervasive paradigm. Frequently, however, it comes in the company of various other approaches, which emerge when the computations that take place on those distributed-memory systems are looked at from different perspectives and at different levels of abstraction.

The remainder of the book is devoted exclusively to message-passing computations over point-to-point connections. Such computations will be described at the task level, which clearly can be regarded as encompassing message-passing computations at the processor level as well. This is so because the latter can be regarded as message-passing computations at the task level when there is exactly one task per processor and two tasks only communicate with each other if they run on processors directly interconnected by a communication link. However, before leaving aside the processor level completely, we find it convenient to have some understanding of how a group of processors interconnected by point-to-point connections can support intertask message passing even among tasks that run on processors not directly connected by a communication link. This is the subject of the following two sections.

1.2 Communication processors

When two tasks that need to communicate with each other run on processors which are not directly interconnected by a communication link, there is no option to perform that intertask communication but to somehow rely on processors other than the two running the tasks to relay the communication traffic as needed. Clearly, then, each processor in the system must, in addition to executing the tasks that run on it, also act as a relayer of the communication traffic that does not originate from (or is destined to) any of the tasks that run on it. Performing this additional function is quite burdensome, so it appears natural to somehow provide the processor with specific capabilities that allow it to do the relaying of communication traffic without interfering with its local computation. In this way, each processor in the system can be viewed as actually a pair of processors that run independently of each other. One of them is the processor that runs the tasks (called the host processor) and the other is the communication processor. Unless confusion may arise, the denomination simply as a processor will in the remainder of the book be used to indicate either the host processor or, as it has been so far, the pair comprising the host processor and the communication processor.

In the context of computer networks (and in a similar fashion networks of workstations as well), the importance of communication processors was recognized at the very beginning, not only by the performance-related reasons we indicated, but mainly because, by the very nature of the services provided by such networks, each communication processor was to provide services to various users at its site. The first generation of distributed-memory multiprocessors, however, was conceived without any concern for this issue, but very soon afterwards it became clear that the communication traffic would be an unsurmountable
bottleneck unless special hardware was provided to handle that traffic. The use of communication processors has been the rule since.

There is a great variety of approaches to the design of a communication processor, and that depends of course on the programming model to be provided at the task level. If message passing is all that needs to be provided, then the communication processor has to at least be able to function as an efficient communication relayer. If, on the other hand, a shared-memory programming model is intended, either by itself or in a hybrid form that also allows message passing, then the communication processor must also be able to handle memory-management functions.

Let us concentrate a little more on the message-passing aspects of communication processors. The most essential function to be performed by a communication processor is in this case to handle the reception of messages, which may come either from the host processor attached to it or from another communication processor, and then to decide where to send it next, which again may be the local host processor or another communication processor. This function per se involves very complex issues, which are the subject of our discussion in Section 1.3.

Another very important aspect in the design of such communication processors comes from viewing them as processors with an instruction set of their own, and then the additional issue comes up of designing such an instruction set so to provide communication services not only to the local host processor but in general to the entire system. The enhanced flexibility that comes from viewing a communication processor in this way is very attractive indeed, and has motivated a few very interesting approaches to the design of those processors. So, for example, in order to send a message to another (remote) task, a task running on the local host processor has to issue an instruction to the communication processor that will tell it to do so. This instruction is the same that the communication processors exchange among themselves in order to have messages passed on as needed until a destination is reached. In addition to rendering the view of how a communication processor handles the traffic of point-to-point messages a little simpler, regarding the communication processor as an instruction-driven entity has many other advantages. For example, a host processor may direct its associated communication processor to perform complex group communication functions and do something else until that function has been completed system-wide. Some very natural candidate functions are discussed in this book, especially in Chapters 4 and 5 (although algorithms presented elsewhere in the book may also be regarded as such, only at a higher level of complexity).

1.3 Routing and flow control

As we remarked in the previous section, one of the most basic and important functions to be performed by a communication processor is to act as a relayer of the messages it receives by either sending them on to its associated host processor or by passing them along to another communication processor. This function is known as routing, and has various important aspects that deserve our attention.

For the remainder of this chapter, we shall let our distributed-memory system be represented by the connected undirected graph $G_p = (N_p,E_p)$, where the set of nodes $N_p$ is the set of processors (each processor viewed as the pair comprising a host processor and a communication processor) and the set $E_p$ of undirected edges is the set of point-to-point bidirectional communication links. A message is normally received at a communication processor as a pair $(q, \text{Msg})$, meaning that $\text{Msg}$ is to be delivered to processor $q$. Here $\text{Msg}$ is the message as it is first issued by the task that sends it, and can be regarded as comprising a pair of fields as well, say $\text{Msg} = (u, \text{msg})$, where $u$ denotes the task running on processor $q$ to which the message is to be delivered and $\text{msg}$ is the message as $u$ must receive it. This implies that at each processor the information of which task runs on which processor must be available, so that intertask messages can be addressed properly when they are first issued. Section 1.6 is devoted to a discussion of how this information can be obtained.

When a processor $r$ receives the message $(q, \text{Msg})$, it checks whether $q = r$ and in the affirmative case forwards $\text{Msg}$ to the host processor at $r$. Otherwise, the message must be destined to another processor, and is then forwarded by the communication processor for
eventual delivery to that other processor. At processor \( r \), this forwarding takes place according to the function \( \text{next}(q) \), which indicates the processor directly connected to \( r \) to which the message must be sent next for eventual delivery to \( q \) (that is, \( (r, \text{next}(q)) \in E_r \)).

The function \( \text{next} \) is a routing function, and ultimately indicates the set of links a message must traverse in order to be transported between any two processors in the system. For processors \( p \) and \( q \), we denote by \( R(p,q) \subseteq E_p \) the set of links to be traversed by a message originally sent by a task running on \( p \) to a task running on \( q \). Clearly, \( R(p,p) = \emptyset \) and in general \( R(p,q) \) and \( R(q,p) \) are different sets.

Routing can be fixed or adaptive, depending on how the function \( \text{next} \) is handled. In the fixed case, the function \( \text{next} \) is time-invariant, whereas in the adaptive case it may be time-varying. Routing can also be deterministic or nondeterministic, depending on how many processors \( \text{next} \) can be chosen from at a processor. In the deterministic case there is only one choice, whereas the nondeterministic case allows multiple choices in the determination of \( \text{next} \).

Pairwise combinations of these types of routing are also allowed, with adaptivity and nondeterminism being usually advocated for increased performance and fault-tolerance.

Advantageous as some of these enhancements to routing may be, not many of adaptive or nondeterministic schemes have made it into practice, and the reason is that many difficulties accompany those enhancements at various levels. For example, the FIFO (First In, First Out) order of message delivery at the processor level cannot be trivially guaranteed in the adaptive or nondeterministic cases, and then so cannot at the task level either, that is, messages sent from one task to another may end up delivered in an order different than the order they were sent. For some applications, as we discuss for example in Section 5.2.1, this would complicate the treatment at the task level and most likely do away with whatever improvement in efficiency one might have obtained with the adaptive or nondeterministic approaches to routing. (We return to the question of ensuring FIFO message delivery among tasks in Section 1.6.2, but in a different context.)

Let us then concentrate on fixed, determinist routing for the remainder of the chapter. In this case, and given a destination processor \( q \), the routing function \( \text{next}(q) \) does not lead to any loops (i.e., by successively moving from processor to processor as dictated by \( \text{next} \) until \( q \) is reached it is not possible to return to an already visited processor). This is so because the existence of such a loop would either require at least two possibilities for the determination of \( \text{next}(q) \) for some \( r \), which is ruled out by the assumption of deterministic routing, or require that \( \text{next} \) be allowed to change with time, which cannot be under the assumption of fixed routing. If routing is deterministic, then another way of arriving at this loopfree property of \( \text{next} \) is to recognize that, for fixed routing, the sets \( R \) of links are such that \( R(r,q) \subseteq R(p,q) \) for every processor \( r \) that can be obtained from \( p \) by successively applying \( \text{next} \) given \( q \). The absence of loops comes as a consequence. Under this alternative view, it becomes clear that, by building the sets \( R \) to contain shortest paths (i.e., paths with the least possible numbers of links) in the fixed, deterministic case, the containments for those sets appear naturally, and then one immediately obtains a routing function with no loops.

Loops in a routing function refer to one single end-to-end directed path (i.e., a sequence of processors obtained by following \( \text{next}(q) \) from \( r = p \) for some \( p \) and fixed \( q \)), and clearly should be avoided. Another related concept, that of a directed cycle in a routing function, can also lead to undesirable behavior in some situations (to be discussed shortly), but cannot be altogether avoided. A directed cycle exists in a routing function when two or more end-to-end directed paths share at least two processors (and sometimes links as well), say \( p \) and \( q \), in such a way that \( q \) can be reached from \( p \) by following \( \text{next}(q) \) at the intermediate \( r \)’s, and so can \( p \) from \( q \) by following \( \text{next}(p) \). Every routing function contains at least the directed cycles implied by the sharing of processors \( p \) and \( q \) by the sets \( R(p,q) \) and \( R(q,p) \) for all \( p,q \in N \). A routing function containing only these directed cycles does not have any end-to-end directed paths sharing links in the same direction, and is referred to as a quasi-acyclic routing function.

Another function that is normally performed by communication processors and goes closely along that of routing is the function of flow control. Once the routing function \( \text{next} \) has been established and the system begins to transport messages among the various pairs of processors, the storage and communication resources that the interconnected communication processors possess must be shared not only by the messages already on
their way to destination processors but also by other messages that continue to be admitted from the host processors. Flow control strategies aim at optimizing the use of the system's resources under such circumstances. We discuss three such strategies in the remainder of this section.

The first mechanism we investigate for flow control is the *store-and-forward* mechanism. This mechanism requires a message \((q, \text{Msg})\) to be divided into *packets* of fixed size. Each packet carries the same addressing information as the original message (i.e., \(q\)), and can therefore be transmitted independently. If these packets cannot be guaranteed to be delivered to \(q\) in the FIFO order, then they must also carry a sequence number, to be used at \(q\) for the re-assembly of the message. (However, guaranteeing the FIFO order is a straightforward matter under the assumption of fixed, deterministic routing, so long as the communication links themselves are FIFO links.) At intermediate communication processors, packets are stored in buffers for later transmission when the required link becomes available (a queue of packets is kept for each link).

Store-and-forward flow control is prone to the occurrence of deadlocks, as the packets compete for shared resources (buffering space at the communication processors, in this case). One simple situation in which this may happen is the following. Consider a cycle of processors in \(G_p\), and suppose that one task running on each of the processors in the cycle has a message to send to another task running on another processor on the cycle that is more than one link away. Suppose in addition that the routing function \(\text{next}\) is such that all the corresponding communication processors, after having received such messages from their associated host processors, attempt to send them in the same direction (clockwise or counterclockwise) on the cycle of processors. If buffering space is no longer available at any of the communication processors on the cycle, then deadlock is certain to occur. This type of deadlock can be prevented by employing what is called a *structured buffer pool*. This is a mechanism whereby the buffers at all communication processors are divided into classes, and whenever a packet is sent between two directly interconnected communication processors, it can only be accepted for storage at the receiving processor if there is buffering space in a specific buffer class, which is normally a function of some of the packet's addressing parameters. If this function allows no cyclic dependency to be formed among the various buffer classes, then deadlock is ensured never to occur. Even with this issue of deadlock resolved, the store-and-forward mechanism suffers from two main drawbacks. One of them is the latency for the delivery of messages, as the packets have to be stored at all intermediate communication processors. The other drawback is the need to use memory bandwidth, which seldom can be provided entirely by the communication processor and has then to be shared with the tasks that run on the associated host processor.

The potentially excessive latency of store-and-forward flow control is partially remedied by the second flow-control mechanism we describe. This mechanism is known as *circuit switching*, and requires an end-to-end directed path to be entirely reserved in one direction for a message before it is transmitted. Once all the links on the path have been secured for that particular transmission, the message is then sent and at the intermediate processors incurs no additional delay waiting for links to become available. The reservation process employed by circuit switching is also prone to the occurrence of deadlocks, as links may participate in several paths in the same direction. Portions of those paths may form directed cycles that may in turn deadlock the reservation of links. Circuit switching should, for this reason, be restricted to those routing functions that are quasi-acyclic, which by definition pose no deadlock threat to the reservation process.

Circuit switching is obviously inefficient for the transmission of short messages, as the time for the entire path to be reserved becomes then prominent. Even for long messages, however, its advantages may not be too pronounced, depending primarily on how the message is transmitted once the links are reserved. If the message is divided into packets that have to be stored at the intermediate communication processors, then the gain with circuit switching may be only marginal, as a packet is only sent on the next link after it has been completely received (all that is saved is then the wait time on outgoing packet queues). It is possible, however, to pipeline the transmission of the message so that only very small portions have to be stored at the intermediate processors, as in the third flow-control strategy we describe next.
The last strategy we describe for flow control employs packet blocking (as opposed to packet buffering or link reservation) as one of its basic paradigms. The resulting mechanism is known as wormhole routing (a misleading denomination, because it really is a flow-control strategy), and contrasting with the previous two strategies, the basic unit on which flow control is performed is not a packet but a flit (flow-control digit). A flit contains no routing information, so every flit in a packet must follow the leading flit, where the routing information is kept when the packet is subdivided. With wormhole routing, the inherent latency of store-and-forward flow control due to the constraint that a packet can only be sent forward after it has been received in its entirety is eliminated. All that needs to be stored is a flit, significantly smaller than a packet, so the transmission of the packet is pipelined, as portions of it may be flowing on different links and portions may be stored. When the leading flit needs access to a resource (memory space or link) that it cannot have immediately, the entire packet is blocked and only proceeds when that flit can advance. As with the previous two mechanisms, deadlock can also arise in wormhole routing. The strategy for dealing with this is to break the directed cycles in the routing function (thereby possibly making pairs of processors inaccessible to each other), then add virtual links to the already existing links in the network, and then finally fix the routing function by the use of the virtual links. Directed cycles in the routing function then become "spirals", and deadlocks can no longer occur. (Virtual links are in the literature referred to as virtual channels, but channels will have in this book a different connotation—cf. Section 1.4.)

In the case of multiprocessors, the use of communication processors employing wormhole routing for flow control tends to be such that the time to transport a message between nodes directly connected by a link in \( G_P \) is only marginally smaller than the time spent when no direct connection exists. In such circumstances, \( G_P \) can often be regarded as being a complete graph (cf. Section 2.1, where we discuss details of the example given in Section 1.6.2).

To finalize this section, we mention that yet another flow-control strategy has been proposed that can be regarded as a hybrid strategy combining store-and-forward flow control and wormhole routing. It is called virtual cut-through, and is characterized by pipelining the transmission of packets as in wormhole routing, and by requiring entire packets to be stored when an outgoing link cannot be immediately used, as in store-and-forward. Virtual cut-through can then be regarded as a variation of wormhole routing in which the pipelining in packet transmission is retained but packet blocking is replaced with packet buffering.

### 1.4 Reactive message-passing programs

So far in this chapter we have discussed how message-passing systems relate to distributed-memory systems, and have outlined some important characteristics at the processor level that allow tasks to communicate with one another by message passing over point-to-point communication channels. Our goal in this section is to introduce, in the form of a template algorithm, our understanding of what a distributed algorithm is and of how it should be described. This template and some of the notation associated with it will in Section 2.1 evolve into the more compact notation that we use throughout the book.

We represent a distributed algorithm by the connected directed graph \( G_T = (N_T, D_T) \), where the node set \( N_T \) is a set of tasks and the set of directed edges \( D_T \) is a set of unidirectional communication channels. (A connected directed graph is a directed graph whose underlying undirected graph is connected.) For a task \( t \), we let \( In_t \subseteq D_T \) denote the set of edges directed towards \( t \) and \( Out_t \subseteq D_T \) the set of edges directed away from \( t \). Channels in \( In_t \) are those on which \( t \) receives messages and channels in \( Out_t \) are those on which \( t \) sends messages. We also let \( n_t = |In_t| \), that is, \( n_t \) denotes the number of channels on which \( t \) may receive messages.

A task \( t \) is a reactive (or message-driven) entity, in the sense that normally it only performs computation (including the sending of messages to other tasks) as a response to the receipt of a message from another task. An exception to this rule is that at least one task must be allowed to send messages out "spontaneously" (i.e., not as a response to a message receipt) to other tasks at the beginning of its execution, inasmuch as otherwise the assumed message-driven character of the tasks would imply that every task would idle indefinitely and no computation would
take place at all. Also, a task may initially perform computation for initialization purposes.

Algorithm \( \text{Task}_t \), given next, describes the overall behavior of a generic task \( t \). Although in this algorithm we (for ease of notation) let tasks compute and then send messages out, no such precedence is in fact needed, as computing and sending messages out may constitute intermingled portions of a task’s actions.

\[\textbf{Algorithm Task}_t:\]

\begin{verbatim}
Do some computation;
send one message on each channel of a (possibly empty) subset of \( \text{Out}_t \);
repeat
  receive message on \( c_1 \in \text{In}_t \text{ and } B_1 \rightarrow \)
  Do some computation;
  send one message on each channel of a (possibly empty) subset of \( \text{Out}_t \)
  or...
  or
  receive message on \( c_{im} \in \text{In}_t \text{ and } B_{im} \rightarrow \)
  Do some computation;
  send one message on each channel of a (possibly empty) subset of \( \text{Out}_t \)
until global termination is known to \( t \).
\end{verbatim}

There are many important observations to be made in connection with Algorithm \( \text{Task}_t \). The first important observation is in connection with how the computation begins and ends for task \( t \). As we remarked earlier, task \( t \) begins by doing some computation and by sending messages to none or more of the tasks to which it is connected in \( G_T \) by an edge directed away from it (messages are sent by means of the operation \textbf{send}). Then \( t \) iterates until a global termination condition is known to it, at which time its computation ends. At each iteration, \( t \) does some computation and may send messages. The issue of global termination will be thoroughly discussed in \textbf{Section 6.2} in a generic setting, and before that in various other chapters it will come up in more particular contexts. For now it suffices to notice that \( t \) acquires the information that it may terminate its local computation by means of messages received during its iterations. If designed correctly, what this information signals to \( t \) is that no message will ever reach it again, and then it may exit the \textbf{repeat...until} loop.

The second important observation is on the construction of the \textbf{repeat...until} loop and on the semantics associated with it. Each iteration of this loop contains \( n_t \) guarded commands grouped together by \textbf{or} connectives. A \textit{guarded command} is usually denoted by

\[ \text{guard} \rightarrow \text{command}, \]

where, in our present context, \textit{guard} is a condition of the form

\[ \text{receive message on } c_k \in \text{In}_t \text{ and } B_k \rightarrow \]

for some Boolean condition \( B_k \), where \( 1 \leq k \leq n_t \). The \textbf{receive} appearing in the description of the \textit{guard} is an operation for a task to receive messages. The \textit{guard} is said to be \textit{ready} when there is a message available for immediate reception on channel \( c_k \) and furthermore the condition \( B_k \) is \textit{true}. This condition may depend on the message that is available for reception, so that a guard may be ready or not, for the same channel, depending on what is at the channel to be received. The overall semantics of the \textbf{repeat...until} loop is then the following. At each iteration, execute the \textit{command} of exactly one guarded command whose \textit{guard} is ready. If no \textit{guard} is ready, then the task is suspended until one is. If more than one \textit{guard} is ready, then one of them is selected arbitrarily. As the reader will verify by our many
distributed algorithm examples along the book, this possibility of nondeterministically selecting guarded commands for execution provides great design flexibility.

Our final important remark in connection with Algorithm \textit{Task_t} is on the semantics associated with the \texttt{receive} and \texttt{send} operations. Although as we have remarked the use of a \texttt{receive} in a \texttt{guard} is to be interpreted as an indication that a message is available for immediate receipt by the task on the channel specified, when used in other contexts this operation in general has a \textit{blocking} nature. A blocking \texttt{receive} has the effect of suspending the task until a message arrives on the channel specified, unless a message is already there to be received, in which case the reception takes place and the task resumes its execution immediately.

The \texttt{send} operation too has a semantics of its own, and in general may be \textit{blocking} or \textit{nonblocking}. If it is blocking, then the task is suspended until the message can be delivered directly to the receiving task, unless the receiving task happens to be already suspended for message reception on the corresponding channel when the \texttt{send} is executed. A blocking \texttt{send} and a blocking \texttt{receive} constitute what is known as \textit{task rendez-vous}, which is a mechanism for task synchronization. If the \texttt{send} operation has a nonblocking nature, then the task transmits the message and immediately resumes its execution. This nonblocking version of \texttt{send} requires buffering for the messages that have been sent but not yet received, that is, messages that are \textit{in transit} on the channel. Blocking and nonblocking \texttt{send} operations are also sometimes referred to as \textit{synchronous} and \textit{asynchronous}, respectively, to emphasize the synchronizing effect they have in the former case. We refrain from using this terminology, however, because in this book the words synchronous and asynchronous will have other meanings throughout (cf. Section 2.1). When used, as in Algorithm \textit{Task_t}, to transmit messages to more than one task, the \texttt{send} operation is assumed to be able to do all such transmissions in parallel.

The relation of blocking and nonblocking \texttt{send} operations with message buffering requirements raises important questions related to the design of distributed algorithms. If, on the one hand, a blocking \texttt{send} requires no message buffering (as the message is passed directly between the synchronized tasks), on the other hand a nonblocking \texttt{send} requires the ability of a channel to buffer an unbounded number of messages. The former scenario poses great difficulties to the program designer, as communication deadlocks occur with great ease when the programming is done with the use of blocking operations only. For this reason, however unreal the requirement of infinitely many buffers may seem, it is customary to start the design of a distributed algorithm by assuming nonblocking operations, and then at a later stage performing changes to yield a program that makes use of the operations provided by the language at hand, possibly of a blocking nature or of a nature that lies somewhere in between the two extremes of blocking and nonblocking \texttt{send} operations.

The use of nonblocking \texttt{send} operations does in general allow the correctness of distributed algorithms to be shown more easily, as well as their properties. We then henceforth assume that, in Algorithm \textit{Task_t}, \texttt{send} operations have a nonblocking nature. Because Algorithm \textit{Task_t} is a template for all the algorithms appearing in the book, the assumption of nonblocking \texttt{send} operations holds throughout.

Another important aspect affecting the design of distributed algorithms is whether the channels in \textit{D_t} deliver messages in the FIFO order or not. Although as we remarked in Section 1.3 this property may at times be essential, we make no assumptions now, and leave its treatment to be done on a case-by-case basis. We do make the point, however, that in the guards of Algorithm \textit{Task_t} at most one message can be available for immediate reception on a FIFO channel, even if other messages have already arrived on that same channel (the available message is the one to have arrived first and not yet received). If the channel is not FIFO, then any message that has arrived can be regarded as being available for immediate reception.
1.5 Handling infinite-capacity channels

As we saw in Section 1.4, the blocking or nonblocking nature of the `send` operations is closely related to the channels ability to buffer messages. Specifically, blocking operations require no buffering at all, while nonblocking operations may require an infinite amount of buffers. Between the two extremes, we say that a channel has capacity $k \geq 0$ if the number of messages it can buffer before either a message is received by the receiving task or the sending task is suspended upon attempting a transmission is $k$. The case of $k = 0$ corresponds to a blocking `send`, and the case in which $k \to \infty$ corresponds to a nonblocking `send`.

Although Algorithm `Task_t` of Section 1.4 is written under the assumption of infinite-capacity channels, such an assumption is unreasonable, and must be dealt with somewhere along the programming process. This is in general achieved along two main steps. First, for each channel $c$ a nonnegative integer $b(c)$ must be determined that reflects the number of buffers actually needed by channel $c$. This number must be selected carefully, as an improper choice may introduce communication deadlocks in the program. Such a deadlock is represented by a directed cycle of tasks, all of which are suspended to send a message on the channel on the cycle, which cannot be done because all channels have been assigned insufficient storage space. Secondly, once the $b(c)$'s have been determined, Algorithm `Task_t` must be changed so that it now employs `send` operations that can deal with the new channel capacities. Depending on the programming language at hand, this can be achieved rather easily. For example, if the programming language offers channels with zero capacity, then each channel $c$ may be replaced with a serial arrangement of $b(c)$ relay tasks alternating with $b(c) + 1$ zero-capacity channels. Each relay task has one input channel and one output channel, and has the sole function of sending on its output channel whatever it receives on its input channel. It has, in addition, a storage capacity of exactly one message, so the entire arrangement can be viewed as a $b(c)$-capacity channel.

The real problem is of course to determine values for the $b(c)$'s in such a way that no new deadlock is introduced in the distributed algorithm (put more optimistically, the task is to ensure the deadlock-freedom of an originally deadlock-free program). In the remainder of this section, we describe solutions to this problem which are based on the availability of a bound $r(c)$, provided for each channel $c$, on the number of messages that may require buffering in $c$ when $c$ has infinite capacity. This number $r(c)$ is the largest number of messages that will ever be in transit on $c$ when the receiving task of $c$ is itself attempting a message transmission, so the messages in transit have to be buffered.

Although determining the $r(c)$'s can be very simple for some distributed algorithms (cf. Sections 5.4 and 8.5), for many others such bounds are either unknown, or known imprecisely, or simply do not exist. In such cases, the value of $r(c)$ should be set to a "large" positive integer $M$ for all channels $c$ whose bounds cannot be determined precisely. Just how large this $M$ has to be, and what the limitations of this approach are, we discuss later in this section.

If the value of $r(c)$ is known precisely for all $c \in D_T$, then obviously the strategy of assigning $b(c) = r(c)$ buffers to every channel $c$ guarantees the introduction of no additional deadlock, as every message ever to be in transit when its destination is engaged in a message transmission will be buffered (there may be more messages in transit, but only when their destination is not engaged in a message transmission, and will therefore be ready for reception within a finite amount of time). The interesting question here is, however, whether it can still be guaranteed that no new deadlock will be introduced if $b(c) < r(c)$ for some channels $c$. This would be an important strategy to deal with the cases in which $r(c) = M$ for some $c \in D_T$, and to allow (potentially) substantial space savings in the process of buffer assignment. Theorem 1.1 given next concerns this issue.

Theorem 1.1

Suppose that the distributed algorithm given by Algorithm `Task_t` for all $t \in N_T$ is deadlock-free. Suppose in addition that $G_T$ contains no directed cycle on which every channel $c$ is such that either $b(c) < r(c)$ or $r(c) = M$. Then the distributed algorithm obtained by replacing each infinite-capacity channel $c$ with a $b(c)$-capacity channel is deadlock-free.
Proof: A necessary condition for a deadlock to arise is that a directed cycle exists in \( G_T \) whose tasks are all suspended on an attempt to send messages on the channels on that cycle. By the hypotheses, however, every directed cycle in \( G_T \) has at least one channel \( c \) for which \( b(c) = r(c) < M \), so at least the tasks \( t \) that have such channels in \( \text{Out} \) are never indefinitely suspended upon attempting to send messages on them. The converse of Theorem 1.1 is also often true, but not in general. Specifically, there may be cases in which \( r(c) = M \) for all the channels \( c \) of a directed cycle, and yet the resulting algorithm is deadlock-free, as \( M \) may be a true upper bound for \( c \) (albeit unknown). So setting \( b(c) = r(c) \) for this channel does not necessarily mean providing it with insufficient buffering space.

As long as we comply with the sufficient condition given by Theorem 1.1, it is then possible to assign to some channels \( c \) fewer buffers than \( r(c) \) and still guarantee that the resulting distributed algorithm is deadlock-free if it was deadlock-free to begin with. In the remainder of this section, we discuss two criteria whereby these channels may be selected. Both criteria lead to intractable optimization problems (i.e., \( NP \)-hard problems), so heuristics need to be devised to approximate solutions to them (some are provided in the literature).

The first criterion attempts to save as much buffering space as possible. It is called the space-optimal criterion, and is based on a choice of \( M \) such that

\[
M > \sum_{c \in D_T - C^+} r(c),
\]

where \( C^+ \) is the set of channels for which a precise upper bound is not known. This criterion requires a subset of channels \( C \subseteq D_T \) to be determined such that every directed cycle in \( G_T \) has at least one channel in \( C \), and such that

\[
\sum_{c \in C} r(c)
\]

is minimum over all such subsets (clearly, \( C \) and \( C^+ \) are then disjoint, given the value of \( M \), unless \( C^+ \) contains the channels of an entire directed cycle from \( G_T \)). Then the strategy is to set

\[
b(c) = \begin{cases} r(c), & \text{if } c \in C; \\ 0, & \text{otherwise}, \end{cases}
\]

which ensures that at least one channel \( c \) from every directed cycle in \( G_T \) is assigned \( b(c) = r(c) \) buffers (Figure 1.1). By Theorem 1.1, this strategy then produces a deadlock-free result if no directed cycle in \( G_T \) has all of its channels in the set \( C^+ \). That this strategy employs the minimum number of buffers comes from the optimal determination of the set \( C \).

The space-optimal approach to buffer assignment has the drawback that the concurrency in intertask communication may be too low, inasmuch as many channels in \( D_T \) may be allocated zero buffers. Extreme situations can happen, as for example the assignment of zero buffers to all the channels of a long directed path in \( G_T \). A scenario might then happen in which all tasks in this path (except the last one) would be suspended to communicate with its successor on the path, and this would only take place for one pair of tasks at a time. When at least one channel \( c \) has insufficient buffers (i.e., \( b(c) < r(c) \)) or is such that \( r(c) = M \), a measure of concurrency that attempts to capture the effect we just described is to take the minimum, over all directed paths in \( G_T \) whose channels \( c \) all have \( b(c) < r(c) \) or \( r(c) = M \), of the ratio

\[
\frac{1}{L + 1}
\]

where \( L \) is the number of channels on the path. Clearly, this measure can be no less than \( 1/|N_T| \) and no more than \( 1/2 \), as long as the assignment of buffers conforms to the hypotheses of Theorem 1.1. The value of \( 1/2 \), in particular, can only be achieved if no directed path with more than one channel exists comprising channels \( c \) such that \( b(c) < r(c) \) or \( r(c) = M \) only. Another criterion for buffer assignment to channels is then the concurrency-optimal criterion, which also seeks to save buffering space, but not to the point
Figure 1.1: A graph $G_T$ is shown in part (a). In the graphs of parts (b) through (d), circular nodes are the nodes of $G_T$, while square nodes represent buffers assigned to the corresponding channel in $G_T$. If $r(c) = 1$ for all $c \in \{c_1, c_2, c_3, c_4\}$, then parts (b) through (d) represent three distinct buffer assignments, all of which deadlock-free. Part (b) shows the strategy of setting $b(c) = r(c)$ for all $c \in \{c_1, c_2, c_3, c_4\}$. Parts (c) and (d) represent, respectively, the results of the space-optimal and the concurrency-optimal strategies.

that the concurrency as we defined might be compromised. This criterion looks for buffer assignments that yield a level of concurrency equal to 1/2, and for this reason does not allow any directed path with more than one channel to have all of its channels assigned insufficient buffers. This alone is, however, insufficient for the value of 1/2 to be attained, as for such it is also necessary that no directed path with more than one channel contain channels $c$ with $r(c) = M$ only. Like the space-optimal criterion, the concurrency-optimal criterion utilizes a value of $M$ such that

$$M > \sum_{c \in D_T - C^+} r(c).$$

This criterion requires a subset of channels $C \subseteq D_T$ to be found such that no directed path with more than one channel exists in $G_T$ comprising channels from $C$ only, and such that

$$\sum_{c \in C} r(c)$$

is maximum over all such subsets (clearly, $C^* \subseteq C$, given the value of $M$, unless $C^*$ contains the channels of an entire directed path from $G_T$ with more than one channel). The strategy is then to set

$$b(c) = \begin{cases} 0, & \text{if } c \in C; \\ r(c), & \text{otherwise}, \end{cases}$$

thereby ensuring that at least one channel $c$ in every directed path with more than one channel in $G_T$ is assigned $b(c) = r(c)$ buffers, and that, as a consequence, at least one channel $c$ from every directed cycle in $G_T$ is assigned $b(c) = r(c)$ buffers as well (Figure 1.1). By Theorem 1.1, this strategy then produces a deadlock-free result if no directed cycle in $G_T$
has all of its channels in the set $C'$. The strategy also provides concurrency equal to 1/2 by our definition, as long as $C'$ does not contain all the channels of any directed path in $G_p$ with more than one channel. Given this constraint that optimal concurrency must be achieved (if possible), then the strategy employs the minimum number of buffers, as the set $C$ is optimally determined.

1.6 Processor allocation

When we discussed the routing of messages among processors in Section 1.3 we saw that addressing a message at the task level requires knowledge by the processor running the task originating the message of the processor on which the destination task runs. This information is provided by what is known as an allocation function, which is a mapping of the form

$$A : N_p \rightarrow N_p$$

where $N_p$ and $N_p$ are, as we recall, the node sets of graphs $G_T$ (introduced in Section 1.4) and $G_R$ (introduced in Section 1.3), respectively. The function $A$ is such that $A(t) = p$ if and only if task $t$ runs on processor $p$.

For many of the systems reviewed in Section 1.1 the allocation function is given naturally by how the various tasks in $N_p$ are distributed throughout the system, as for example computer networks and networks of workstations. However, for multiprocessors and also for networks of workstations when viewed as parallel processing systems, the function $A$ has to be determined during what is called the processor allocation step of program design. In these cases, $G_T$ should be viewed not simply as the task graph introduced earlier, but rather as an enlargement of that graph to accommodate the relay tasks discussed in Section 1.4 (or any other tasks with similar functions—cf. Exercise 4).

The determination of the allocation function $A$ is based on a series of attributes associated with both $G_T$ and $G_R$. Among the attributes associated with $G_T$ is its routing function, which, as we remarked in section 1.3, can be described by the mapping

$$R : N_p \times N_p \rightarrow \mathbb{P}$$

For all $p,q \in N_p, R(p,q)$ is the set of links on the route from processor $p$ to processor $q$,

possibly distinct from $R(q,p)$ and such that $R(p,p) = \emptyset$. Additional attributes of $G_T$ are the relative processor speed (in instructions per unit time) of $p \in N_p$, $s_p$, and the relative link capacity (in bits per unit time) of $(p,q) \in E_p, c_{p,q}$ (the same in both directions). These numbers are such that the ratio $s_p/s_q$ indicates how faster processor $p$ is than processor $q$; similarly for the communication links.

The attributes of graph $G_T$ are the following. Each task $t$ is represented by a relative processing demand (in number of instructions) $\psi_t$, while each channel $(t \rightarrow u)$ is represented by a relative communication demand (in number of bits) from task $t$ to task $u$, $\zeta(t \rightarrow u)$, possibly different from $\zeta(u \rightarrow t)$The ratio $\psi_t/\psi_u$ is again indicative of how much more processing task $t$ requires than task $u$, the same holding for the communication requirements.

The process of processor allocation is generally viewed as one of two main possibilities. It may be static, if the allocation function $A$ is determined prior to the beginning of the computation and kept unchanged for its entire duration, or it may be dynamic, if $A$ is allowed to change during the course of the computation. The former approach is suitable to cases in which both $G_T$ and $G_R$, as well as their attributes, vary negligibly with time. The dynamic approach, on the other hand, is more appropriate to cases in which either the graphs or their attributes are time-varying, and then provides opportunities for the allocation function to be revised in the light of such changes. What we discuss in Section 1.6.1 is the static allocation of processors to tasks. The dynamic case is usually much more difficult, as it requires tasks to be migrated among processors, thereby interfering with the ongoing computation.

Successful results of such dynamic approaches are for this reason scarce, except for some attempts that can in fact be regarded as a periodic repetition of the calculations for static processor allocation, whose resulting allocation functions are then kept unchanged for the duration of the period. We do nevertheless address the question of task migration in Section 1.6.2 in the context of ensuring the FIFO delivery of messages among tasks under such circumstances.
1.6.1 The static approach

The quality of an allocation function \( A \) is normally measured by a function that expresses the time for completion of the entire computation, or some function of this time. This criterion is not accepted as a consensus, but it seems to be consonant with the overall goal of parallel processing systems, namely to compute faster. So obtaining an allocation function by the minimization of such a function is what one should seek. The function we utilize in this book to evaluate the efficacy of an allocation function \( A \) is the function \( H(A) \) given by

\[
H(A) = \alpha H_P(A) + (1 - \alpha) H_C(A),
\]

where \( H_P(A) \) gives the time spent with computation when \( A \) is followed, \( H_C(A) \) gives the time spent with communication when \( A \) is followed, and \( \alpha \) such that \( 0 < \alpha < 1 \) regulates the relative importance of \( H_P(A) \) and \( H_C(A) \). This parameter \( \alpha \) is crucial, for example, in conveying to the processor allocation process some information on how efficient the routing mechanisms for interprocessor communication are (cf. Section 1.3).

The two components of \( H(A) \) are given respectively by

\[
H_P(A) = \sum_{t} \psi_t \frac{s_t}{s_p}
\]

and

\[
H_C(A) = \sum_{(p,q) \in E_P} \sum_{(t \rightarrow u) \in E_T} \frac{1}{c(p,q)} \zeta(t \rightarrow u).
\]

This definition of \( H_P(A) \) has two types of components. One of them, \( \psi_t s_t \), accounts for the time to execute task \( t \) on processor \( p \). The other component, \( \psi_t \psi_u s_p \), is a function of the additional time incurred by processor \( p \) when executing both tasks \( t \) and \( u \) (various other functions can be used here, as long as nonnegative). If an allocation function \( A \) is sought by simply minimizing \( H_P(A) \) then the first component will tend to lead to an allocation of the fastest processors to run all tasks, while the second component will lead to a dispersion of the tasks among the processors. The definition of \( H_C(A) \), in turn, embodies components of the type \( \zeta(t \rightarrow u) c(p,q) \), which reflects the time spent in communication from task \( t \) to task \( u \) on link \( (p,q) \in R(A(t), A(u)) \). Contrasting with \( H_P(A) \), if an allocation function \( A \) is sought by simply minimizing \( H_C(A) \), then tasks will tend to be concentrated on a few processors. The minimization of the overall \( H(A) \) is then an attempt to reconcile conflicting goals, as each of its two components tend to favor different aspects of the final allocation function.

As an example, consider the two-processor system comprising processors \( p \) and \( q \). Consider also the two tasks \( t \) and \( u \). If the allocation function \( A_1 \) assigns \( p \) to run \( t \) and \( q \) to run \( u \), then we have. assuming \( \alpha = 1/2 \),

\[
2H(A_1) = \frac{\psi_t}{s_p} + \frac{\psi_u}{s_q} + \frac{\zeta(t \rightarrow u) + \zeta(u \rightarrow t)}{c(p,q)}.
\]

An allocation function \( A_2 \) assigning \( p \) to run both \( t \) and \( u \) yields

\[
2H(A_2) = \frac{\psi_t}{s_p} + \frac{\psi_u}{s_p} + \frac{\psi_t \psi_u}{s_p}.
\]

Clearly, the choice between \( A_1 \) and \( A_2 \) depends on how the system's parameters relate to one another. For example, if \( s_p = s_q \), then \( A_1 \) is preferable if the additional cost of processing the two tasks on \( p \) is higher than the cost of communication between them over the link \( (p,q) \), that is, if

\[
\frac{\psi_t \psi_u}{s_p} > \frac{\zeta(t \rightarrow u) + \zeta(u \rightarrow t)}{c(p,q)}.
\]

Finding an allocation function \( A \) that minimizes \( H(A) \) is a very difficult problem, \( NP \)-hard in fact, as the problems we encountered in Section 1.5. Given this inherent difficulty, all that is left is to resort to heuristics that allow a "satisfactory" allocation function to be found, that is, an allocation function that can be found reasonably fast and that does not lead to a poor performance of the program. The reader should refer to more specialized literature for various such heuristics.
1.6.2 Task migration
As we remarked earlier in Section 1.6, the need to migrate tasks from one processor to another arises when a dynamic processor allocation scheme is adopted. When tasks migrate, the allocation function \( A \) has to be updated throughout all those processors running tasks that may send messages, according to the structure of \( G_T \), to the migrating task. While performing such an update may be achieved fairly simply (cf. the algorithms given in Section 4.1), things become more complicated when we add the requirement that messages continue to be delivered in the FIFO order. We are in this section motivated not only by the importance of the FIFO property in some situations, as we mentioned earlier, but also because solving this problem provides an opportunity to introduce a nontrivial, yet simple, distributed algorithm at this stage in the book. Before we proceed, it is very important to make the following observation right away. The distributed algorithm we describe in this section is not described by the graph \( G_T \), but rather uses that graph as some sort of a “data structure” to work on. The graph on which the computation actually takes place is a task graph having exactly one task for each processor and two unidirectional communication channels (one in each direction) for every two processors in the system. It is then a complete undirected graph or node set \( N_p \), and for this reason we describe the algorithm as if it were executed by the processors themselves. Another important observation, now in connection with \( G_T \), is that its links are assumed to deliver interprocessor messages in the FIFO order (otherwise it would be considerably harder to attempt this at the task level). The reader should notice that considering a complete undirected graph is a means of not having to deal with the routing function associated with \( G_p \) explicitly, which would be necessary if we described the algorithm for \( G_p \).

The approach we take is based on the following observation. Suppose for a moment and for simplicity that tasks are not allowed to migrate to processors where they have already been. and consider two tasks \( u \) and \( v \) running respectively on processors \( p \) and \( q \). If \( v \) migrates to another processor, say \( q' \), and \( p \) keeps sending to processor \( q \) all of \( u \)'s messages destined to task \( v \), and in addition processor \( q \) forwards to processor \( q' \) whatever messages it receives destined to \( v \), then the desired FIFO property is maintained. Likewise, if \( u \) migrates to another processor, say \( p' \), and every message sent by \( u \) is routed through \( p \) first, then the FIFO property is maintained as well. If later these tasks migrate to yet other processors, then the same forwarding scheme still suffices to maintain the FIFO order. Clearly, this scheme cannot be expected to support any efficient computation, as messages tend to follow ever longer paths before eventual delivery. However, this observation serves the purpose of highlighting the presence of a line of processors that initially contains two processors (\( p \) and \( q \)) and increases with the addition of other processors (\( p' \) and \( q' \) being the first) as \( u \) and \( v \) migrate. What the algorithm we are about to describe does, while allowing tasks to migrate even to processors where they ran previously, is to shorten this line whenever a task migrates out of a processor by removing that processor from the line. We call such a line a pipe to emphasize the FIFO order followed by messages sent along it, and for tasks \( u \) and \( v \) denote it by \( \text{pipe}(u, v) \).

This pipe is a sequence of processors sharing the property of running (or having run) at least one of \( u \) and \( v \). In addition, \( u \) runs on the first processor of the pipe, and \( v \) on the last processor. When \( u \) or \( v \) (or both) migrates to another processor, thereby stretching the pipe, the algorithm we describe in the sequel removes from the pipe the processor (or processors) where the task (or tasks) that migrated ran. Adjacent processors in a pipe are not necessarily connected by a communication link in \( G_p \), and in the beginning of the computation the pipe contains at most two processors.

A processor \( p \) maintains, for every task \( u \) that runs on it and every other task \( v \) such that \((u \rightarrow v) \in \text{Out}_u\), a variable \( \text{pipe}_p(u, v) \) to store its view of \( \text{pipe}(u, v) \). Initialization of this variable must be consonant with the initial allocation function. In addition, for every task \( v \), at \( p \) the value of \( A(v) \) is only an indication of the processor on which task \( v \) is believed to run, and is therefore denoted more consistently by \( A_p(v) \). It is to \( A_p(v) \) that messages sent to \( v \) by other tasks running on \( p \) get sent. Messages destined to \( v \) that arrive at \( p \) after \( v \) has migrated out of \( p \) are also sent to \( A_p(v) \). A noteworthy relationship at \( p \) is the following. If \( v \in \text{Out}_u \), then \( \text{pipe}_p(u, v) = <p, \ldots, q> \) if and only if \( A_p(v) = q \). Messages sent to \( A_p(v) \) are then actually being sent on \( \text{pipe}(u, v) \).
First we informally describe the algorithm for the single pipe \( \text{pipe}(u, v) \), letting \( p \) be the processor on which \( u \) runs (i.e., the first processor in the pipe) and \( q \) the processor on which \( v \) runs (i.e., the last processor in the pipe). The essential idea of the algorithm is the following. When \( u \) migrates from \( p \) to another processor \( p' \), processor \( p \) sends a message \( \text{flush}(u, v, p') \) along \( \text{pipe}_p(u, v) \). This message is aimed at informing processor \( q \) (or processor \( q' \), to which task \( v \) may have already migrated) that \( u \) now runs on \( p' \), and also "pushes" every message still in transit from \( u \) to \( v \) along the pipe (it flushes the pipe). When this message arrives at \( q \) (or \( q' \) the pipe is empty and \( A_q(u) \) (or \( A_q(u) \)) may then be updated. A message \( \text{flushed}(u, v, q) \) (or \( \text{flushed}(u, v, q') \)) is then sent directly to \( p' \), which then updates \( A_q(v) \) and its view of the pipe by altering the contents of \( \text{pipe}_q(u, v) \). Throughout the entire process, task \( u \) is suspended, and as such does not compute or migrate.

![Figure 1.2](image)

**Figure 1.2:** When task \( u \) migrates from processor \( p \) to processor \( p' \) and \( v \) from \( q \) to \( q' \), a flush(u, v, p') message and a flush-request(u, v) message are sent concurrently, respectively by \( p \) to \( q \) and by \( q \) to \( p \). The flush message gets forwarded by \( q \) to \( q' \), and eventually causes \( q' \) to send \( p' \) a flushed(u, v, q') message.

This algorithm may also be initiated by \( q \) upon the migration of \( v \) to \( q' \), and then \( v \) must also be suspended. In this case, a message \( \text{flush\_request}(u, v) \) is sent by \( q \) to \( p \), which then engages in the flushing procedure we described after suspending task \( u \). There is also the possibility that both \( p \) and \( q \) initiate concurrently. This happens when \( u \) and \( v \) both migrate (to \( p' \) and \( q' \), respectively) concurrently, i.e., before news of the other task's migration is received. The procedures are exactly the same, with only the need to ensure that \( \text{flush}(u, v, p') \) is not sent again upon receipt of a \( \text{flush\_request}(u, v) \), as it must already have been sent (Figure 1.2).

When a task \( u \) migrates from \( p \) to \( p' \), the procedure we just described is executed concurrently for every \( \text{pipe}(u, v) \) such that \( (u \rightarrow v) \in \text{Out}_p \) and every \( \text{pipe}(v, u) \) such that \( (v \rightarrow u) \in \text{In}_v \). Task \( u \) may only resume its execution at \( p' \) (and then possibly migrate once again) after all the pipes \( \text{pipe}(u, v) \) such that \( (u \rightarrow v) \in \text{Out}_p \) and \( \text{pipe}(v, u) \) such that \( (v \rightarrow u) \in \text{In}_v \) have been flushed, and is then said to be active (it is inactive otherwise, and may not migrate). Task \( u \) also becomes inactive upon the receipt of a \( \text{flush\_request}(u, v) \) when running on \( p \). In this case, only after \( \text{pipe}_p(u, v) \) is updated can \( u \) become once again active. Later in the book we return to this algorithm, both to provide a more formal description of it (in Section 2.1), and to describe its correctness and complexity properties (in Section 2.1 and Section 3.2.1).

### 1.7 Remarks on program development

The material presented in Sections 1.4 through 1.6 touches various of the fundamental issues involved in the design of message-passing programs, especially in the context of multiprocessors, where the issues of allocating buffers to communication channels and processors to tasks are most relevant. Of course not always does the programmer have full access to or control of such issues, which are sometimes too tightly connected to built-in characteristics of the operating system or the programming language, but some level of awareness of what is really happening can only be beneficial.

Even when full control is possible, the directions provided in the previous two sections should not be taken as much more than that. The problems involved in both sections are, as we
mentioned, probably intractable from the standpoint of computational complexity, so that the optima that they require are not really achievable. Also the formulations of those problems can be in many cases troublesome, because they involve parameters whose determination is far from trivial, like for example the upper bound $M$ used in Section 1.5 to indicate our inability in determining tighter values, or the $\alpha$ used in Section 1.6 to weigh the relative importance of computation versus communication in the function $H$. This function cannot be trusted too blindly either. because there is no assurance that, even if the allocation that optimizes it could be found efficiently, no other allocation would in practice provide better results albeit its higher value for $H$.

Imprecise and troublesome though they may be, the guidelines given in Sections 1.5 and 1.6 do nevertheless provide a conceptual framework within which one may work given the constraints of the practical situation at hand. In addition, they in a way bridge the abstract description of a distributed algorithm we gave in Section 1.4 to what tends to occur in practice.

1.8 Exercises

1. For $d \geq 0$, a $d$-dimensional hypercube is an undirected graph with $2^d$ nodes in which every node has exactly $d$ neighbors. If nodes are numbered from 0 to $2^d - 1$, then two nodes are neighbors if and only if the binary representations of their numbers differ by exactly one bit. One routing function that can be used when GP is a hypercube is based on comparing the number of a message’s destination processor, say $q$, with the number of the processor where the message is, say $r$. The message is forwarded to the neighbor of $r$ whose number differs from that of $r$ in the least-significant bit at which the numbers of $q$ and $r$ differ. Show that this routing function is quasi-acyclic.

2. In the context of Exercise 1, consider the use of a structured buffer pool to prevent deadlocks when flow control is done by the store-and-forward mechanism. Give details of how the pool is to be employed for deadlock prevention. How many buffer classes are required?

3. In the context of Exercise 1, explain in detail why the reservation of links when doing flow control by circuit switching is deadlock-free.

4. Describe how to obtain channels with positive capacity from zero-capacity channels, under the constraint the exactly two additional tasks are to be employed per channel of GT.

1.9 Bibliographic notes

Sources in the literature to complement the material of Section 1.1 could hardly be more plentiful. For material on computer networks, the reader is referred to the traditional texts by
Bertsekas and Gallager (1987) and by Tanenbaum (1988), as well as to more recent material on the various aspects of ATM networks (Bae and Suda, 1991; Stamoulis, Anagnostou, and Georgantas, 1994). Networks of workstations are also well represented by surveys (e.g., Bernard, Steve, and Simatic, 1993), as well as by more specific material (Blumofe and Park, 1994).

References on multiprocessors also abound, ranging from reports on early experiences with shared-memory (Gehringer, Siewiorek, and Segall, 1987) and message-passing systems (Hillis, 1985; Seitz, 1985; Arlauskas, 1988; Grunwald and Reed, 1988; Pase and Larrabee, 1988) to the more recent revival of distributed-memory architectures that provide a shared address space (Fernandes, de Amorim, Barbosa, França, and de Souza, 1989; Martonosi and Gupta, 1989; Bell, 1992; Bagheri, Ilin, and Ridgeway Scott, 1994; Reinhardt, Larus, and Wood, 1994; Protic, Tomašević, and Milutinović, 1995). The reader of this book may be particularly interested in the recent recognition that explicit message-passing is often needed, and in the resulting architectural proposals, as for example those of Kranz, Johnson, Agarwal, Kubitowicz, and Lim (1993), Kuskin, Ofelt, Heinrich, Heinlein, Simoni, Gharachorloo, Chapin, Nakahira, Baxter, Horowitz, Gupta, Rosenblum, and Hennessy (1994), Heinlein, Gharachorloo, Dresser, and Gupta (1994), Heinrich, Kuskin, Ofelt, Heinlein, Singh, Simoni, Gharachorloo, Baxter, Nakahira, Horowitz, Gupta, Rosenblum, and Hennessy (1994), and Agarwal, Bianchini, Chaiken, Johnson, Kranz, Kubitowicz, Lim, Mackenzie, and Yeung (1995). Pertinent theoretical insights have also been pursued (Bar-Noy and Dolev, 1993).

The material in Section 1.2 can be expanded by referring to a number of sources in which communication processors are discussed. These include, for example, Dally, Chao, Chien, Hassoun, Horwat, Kaplan, Song, Toty, and Wills (1987), Ramachandran, Solomon, and Vernon (1987), Barbosa and França (1988), and Dally (1990). The material in Barbosa and França (1988) is presented in considerably more detail by Drummond (1990), and, in addition, has pioneered the introduction of messages as instructions to be performed by communication processors. These were later re-introduced under the denomination of active messages (von Eicken, Culler, Goldberg, and Schausier, 1992; Tucker and Mainwaring, 1994).

In addition to the aforementioned classic sources on computer networks, various other references can be looked up to complement the material on routing and flow control discussed in Section 1.3. For example, the original source for virtual cut-through is Kermani and Kleinrock (1979), while Günther (1981) discusses techniques for deadlock prevention in the store-and-forward case and Gerla and Kleinrock (1982) provide a survey of early techniques. The original publication on wormhole routing is Dally and Seitz (1987), and Gaughan and Yalamanchili (1993) should be looked up by those interested in adaptive techniques. Wormhole routing is also surveyed by Ni and McKinley (1993), and Awerbuch, Kutten, and Peleg (1994) return to the subject of deadlock prevention in the store-and-forward case.

The template given by Algorithm Task_t of Section 1.4 originates from Barbosa (1990a), and the concept of a guarded command on which it is based dates back to Dijkstra (1975). The reader who wants a deeper understanding of how communication channels of zero and nonzero capacities relate to each other may wish to check Barbosa (1990b), which contains a mathematical treatment of concurrency-related concepts associated with such capacities. What this work does is to start at the intuitive notion that greater channel capacity leads to greater concurrency (present, for example, in Gentleman (1981)), and then employ (rather involved) combinatorial concepts related to the coloring of graph edges (Edmonds, 1965; Fulkerson, 1972; Fiorini and Wilson, 1977; Stahl, 1979) to argue that such a notion may not be correct. The Communicating Sequential Processes (CSP) introduced by Hoare (1978) constitute an example of notation based on zero-capacity communication.

Section 1.5 is based on Barbosa (1990a), where in addition a heuristic is presented to support the concurrency-optimal criterion for buffer assignment to channels. This heuristic employs an algorithm to find maximum matchings in graphs (Syslo, Deo, and Kowalik, 1983). The reader has many options to complement the material of Section 1.6. References on the intractability of processor allocation (in the sense of NP-hardness, as in Karp (1972) and Garey and Johnson (1979)) are Krumme, Venkataraman, and Cybenko (1986) and Ali and
El-Rewini (1994). For the static approach, some references are Ma, Lee, and Tsuchiya (1982), Shen and Tsai (1985), Sinclair (1987), Barbosa and Huang (1988)—on which Section 1.6.1 is based, Ali and El-Rewini (1993), and Selvakumar and Siva Ram Murthy (1994). The material in Barbosa and Huang (1988) includes heuristics to overcome intractability that are based on neural networks (as is the work of Fox and Furmanski (1988)) and on the $A^*$ algorithm for heuristic search (Nilsson, 1980; Pearl, 1984). A parallel variation of the latter algorithm (Freitas and Barbosa, 1991) can also be employed. Fox, Kolawa, and Williams (1987) and Nicol and Reynolds (1990) offer treatments of the dynamic type. References on task migration include Theimer, Lantz, and Cheriton (1985), Ousterhout, Cherenson, Douglis, Nelson, and Welch (1988), Ravi and Jefferson (1988), Eskicioğlu and Cabrera (1991), and Barbosa and Porto (1995)—which is the basis for our treatment in Section 1.6.2.

Details on the material discussed in Section 1.7 can be found in Hellmuth (1991), or in the more compact accounts by Barbosa, Drummond, and Hellmuth (1991a; 1991b; 1994).


Chapter 2: **Intrinsic Constraints**

**Overview**

This chapter, like Chapter 1, still has the flavor of a chapter on preliminaries, although various distributed algorithms are presented and analyzed in its sections. The reason why it is still in a way a chapter on preliminary concepts is that it deals mostly with constraints on the computations that may be carried out over the model introduced in Section 1.4 for distributed computations by point-to-point message passing.

Initially, in Section 2.1, we return to the graph-theoretic model of Section 1.4 to specify two of the variants that it admits when we consider its timing characteristics. These are the fully asynchronous and fully synchronous variants that will accompany us throughout the book. For each of the two, Section 2.1 contains an algorithm template, which again is used through the remaining chapters. In addition to these templates, in Section 2.1 we return to the problem of ensuring the FIFO delivery of intertask messages when tasks migrate discussed in Section 1.6.2. The algorithm sketched in that section to solve the problem is presented in full in Section 2.1 to illustrate the notational conventions adopted for the book. In addition, once the algorithm is known in detail, some of its properties, including some complexity-related ones, are discussed.

Sections 2.2 and 2.3 are the sections in which some of our model's intrinsic constraints are discussed. The discussion in Section 2.2 is centered on the issue of anonymous systems, and in this context several impossibility results are presented. Along with these impossibility results, distributed algorithms for the computations that can be carried out are given and to some extent analyzed.

In Section 2.3 we present a somewhat informal discussion of how various notions of knowledge translate into a distributed algorithm setting, and discuss some impossibility results as well. Our approach in this section is far less formal and complete than in the rest of the book because the required background for such a complete treatment is normally way outside what is expected of this book's intended audience. Nevertheless, the treatment we offer is intended to build up a certain amount of intuition, and at times in the remaining chapters we return to the issues considered in Section 2.3.

Exercises and bibliographic notes follow respectively in Sections 2.4 and 2.5.

### 2.1 Full asynchronism and full synchronism
We start by recalling the graph-theoretic model introduced in Section 1.4, according to which a distributed algorithm is represented by the connected directed graph $G_t = (N_t, D_t)$. In this graph, $N_t$ is the set of tasks and $D_t$ is the set of unidirectional communication channels. Tasks in $N_t$ are message-driven entities whose behavior is generically depicted by Algorithm $Task_t$ (cf. Section 1.4), and the channels in $D_t$ are assumed to have infinite capacity, i.e., no task is ever suspended upon attempting to send a message on a channel (reconciling this assumption with the reality of practical situations was our subject in Section 1.5). Channels in $D_t$ are not generally assumed to be FIFO channels unless explicitly stated.

For the remainder of the book, we simplify our notation for this model in the following manner. The graph $G_t = (N_t, D_t)$ is henceforth denoted simply by $G = (N, D)$, with $n = |N|$ and $m = |D|$. For $1 \leq i, j \leq n$, $n_i$ denotes a member of $N$, referred to simply as a node, and if $j \neq i$ we let $(n_i \rightarrow n_j)$ denote a member of $D$, referred to simply as a directed edge (or an edge, if confusion may not arise). The set of edges directed away from $n_i$ is denoted by $Out \subseteq D$, and the set of edges directed towards $n_i$ is denoted by $In \subseteq D$. Clearly, $(n_i \rightarrow n_j) \in Out$ if and only if $(n_i, n_j) \in E$. The nodes $n_i$ and $n_j$ are said to be neighbors of each other if and only if either $(n_i \rightarrow n_j) \in D$ or $(n_j \rightarrow n_i) \in D$. The set of $n_i$'s neighbors is denoted by $Neig$, and contains two partitions, $I_Neig$ and $O_Neig$, whose members are respectively $n_i$'s neighbors $n_j$ such that $(n_i \rightarrow n_j) \in D$ and $n_i$ such that $(n_i \rightarrow n_j) \in D$. Often $G$ is such that $(n_i \rightarrow n_j) \in D$ if and only if $(n_i, n_j) \in E$, and in this case viewing these two directed edges as the single undirected edge $(n_i, n_j)$ is more convenient. In this undirected case, $G$ is denoted by $G = (N, E)$, and then $m = |E|$. Members of $E$ are referred to simply as edges. In the undirected case, the set of edges incident to $n_i$ is denoted by $Inc \subseteq E$. Two nodes $n_i$ and $n_j$ are neighbors if and only if $(n_i, n_j) \in E$. The set of $n_i$'s neighbors continues to be denoted by $Neig$.

Our main concern in this section is to investigate the nature of the computations carried out by $G$'s nodes with respect to their timing characteristics. This investigation will enable us to complete the model of computation given by $G$ with the addition of its timing properties. The first model we introduce is the fully asynchronous (or simply asynchronous) model, which is characterized by the following two properties.

- Each node is driven by its own, local, independent time basis, referred to as its local clock.
- The delay that a message suffers to be delivered between neighbors is finite but unpredictable.

The complete asynchronism assumed in this model makes it very realistic from the standpoint of somehow reflecting some of the characteristics of the systems discussed in Section 1.1. It is this same asynchronism, however, that accounts for most of the difficulties encountered during the design of distributed algorithms under the asynchronous model. For this reason, frequently a far less realistic model is used, one in which $G$’s timing characteristics are pushed to the opposing extreme of complete synchronism. We return to this other model later in this section.

One important fact to notice is that the notation used to describe a node’s computation in Algorithm $Task_t$ (cf. Section 1.4) is quite well suited to the assumptions of the asynchronous model, because in that algorithm, except possibly initially, computation may only take place at the reception of messages, which are in turn accepted nondeterministically when there is more than one message to choose from. In addition, no explicit use of any timing information is made in Algorithm $Task_t$ (although the use of timing information drawn from the node's local clock would be completely legitimate and in accordance with the assumptions of the model). According to Algorithm $Task_t$, the computation of a node in the asynchronous model can be described by providing the actions to be taken initially (if that node is to start its computation and send messages spontaneously, as opposed to doing it
in the wake of the reception of a message) and the actions to be taken upon receiving messages when certain Boolean conditions hold. Such a description is given by Algorithm _A_Template_, which is a template for all the algorithms studied in this book under the asynchronous model, henceforth referred to as _asynchronous algorithms_. Algorithm _A_Template_ describes the computation carried out by a node in \( N \). In this algorithm, and henceforth, we let \( N_0 \subseteq N \) denote the nonempty set of nodes that may send messages spontaneously. The prefix _A_ in the algorithm's denomination is meant to indicate that it is asynchronous, and is used in the names of all the asynchronous algorithms in the book.

Algorithm _A_Template_ is given for the case in which \( G \) is a directed graph. For the undirected case, all that needs to be done to the algorithm is to replace all occurrences of both \( In \) and \( Out \) with \( Inc \).

**Algorithm A_Template:**

| Variables: |
| Variables used by \( n \), and their initial values, are listed here. |

**Listing 2.1**

| Input: |
| \( msg = \text{nil} \). |

| Action if \( n \in N_0 \): |
| Do some computation; |
| Send one message on each edge of a (possibly empty) subset of \( Out \). |

**Listing 2.2**

| Input: |
| \( msg \) such that \( \text{origin}(msg) = c_k \in In \), with \( 1 \leq k \leq |In| \). |

| Action when \( B_k \): |
| Do some computation; |
| Send one message on each edge of a (possibly empty) subset of \( Out \). |

Before we proceed to an example of how a distributed algorithm can be expressed according to this template, there are some important observations to make in connection with Algorithm _A_Template_. The first observation is that the algorithm is given by listing the variables it employs (along with their initial values) and then a series of _input/action pairs_. Each of these pairs, in contrast with Algorithm _Task_t_, is given for a specific message type, and may then correspond to more than one guarded command in Algorithm _Task_t_ of Section 1.4, with the input corresponding to the message reception in the _guard_ and the action corresponding to the _command_ part, to be executed when the Boolean condition expressed in the _guard_ is _true_. Conversely, each guarded command in Algorithm _Task_t_ may also correspond to more than one _input/action pair_ in Algorithm _A_Template_. In addition, in order to preserve the functioning of Algorithm _Task_t_, namely that a new
guarded command is only considered for execution in the next iteration, therefore after the command in the currently selected guarded command has been executed to completion, each action in Algorithm A_Template is assumed to be an atomic action. An atomic action is an action that is allowed to be carried out to completion before any interrupt. All actions are numbered to facilitate the discussion of the algorithm's properties.

Secondly, we make the observation that the message associated with an input, denoted by \(msg\), is if \(n \in N_0\) treated as if \(msg = \text{nil}\), since in such cases no message really exists to trigger \(n\)'s action, as in (2.1). When a message does exist, as in (2.2), we assume that its origin, in the form of the edge on which it was received, is known to \(n\). Such an edge is denoted by \(\text{origin}(msg) \in I_n\). In many cases, knowing the edge \(\text{origin}(msg)\) can be regarded as equivalent to knowing \(n_i \in \text{I-Neig}\), for \(\text{origin}(msg) = (n_i \rightarrow n)\) (that is, \(n_i\) is the node from which \(msg\) originated). Similarly, sending a message on an edge in \(\text{Out}\) is in many cases equivalent to sending a message to \(n_i \in \text{O-Neig}\) if that edge is \((n_i \rightarrow n_i)\). However, we refrain from stating these as general assumptions because they do not hold in the case of anonymous systems, treated in Section 2.2. When they do hold and \(G\) is an undirected graph, then all occurrences of \(\text{I-Neig}\) and of \(\text{O-Neig}\) in the modified Algorithm A_Template must be replaced with occurrences of \(\text{Neig}\).

As a final observation, we recall that, as in the case of Algorithm Task_t, whenever in Algorithm A_Template \(n\) sends messages on a subset of \(\text{Out}\), containing more than one edge, it is assumed that all such messages may be sent in parallel.

We now turn once again to the material introduced in Section 1.6.2, namely a distributed algorithm to ensure the FIFO order of message delivery among tasks that migrate from processor to processor. As we mentioned in that section, this is an algorithm described on a complete undirected graph that has a node for every processor. So for the discussion of this algorithm \(G\) is the undirected graph \(G = (N, E)\). We also mentioned in Section 1.6.2 that the directed graph whose nodes represent the migrating tasks and whose edges represent communication channels is in this algorithm used as a data structure. While treating this problem, we then let this latter graph be denoted, as in Section 1.6.2, by \(G_r = (N_r, D_r)\), along with the exact same notation used in that section with respect to \(G\). Care should be taken to avoid mistaking this graph for the directed version of \(G\) introduced at the beginning of this section.

Before introducing the additional notation that we need, let us recall some of the notation introduced in Section 1.6.2. Let \(A\) be the initial allocation function. For a node \(n\) and every task \(u\) such that \(A(u) = n\), a variable \(\text{pipe}(u, v)\) for every task \(v\) such that \((u \rightarrow v) \in \text{Out}\), indicates \(n\)'s view of \(\text{pipe}(u, v)\). Initially, \(\text{pipe}(u, v) = n\). In addition, for every task \(v\) a variable \(A(v)\) is used by \(n\) to indicate the node where task \(v\) is believed to run. This variable is initialized such that \(A(v) = A(v)\). Messages arriving at \(n\) destined to \(v\) are assumed to be sent to \(A(v)\) if \(A(v) \neq n\), or to be kept in a FIFO queue, called \(\text{queue}\), otherwise.

Variables employed in connection with task \(u\) are the following. The Boolean variable \(\text{active}_u\) (initially set to \(\text{true}\)) is used to indicate whether task \(u\) is active. Two counters, \(\text{pending}_{\text{in}}\) and \(\text{pending}_{\text{out}}\), are used to register the number of pipes that need to be flushed before \(u\) can once again become active. The former counter refers to pipes \(\text{pipe}(v, u)\) such that \((v \rightarrow u) \in \text{In}_v\), and the latter to pipes \(\text{pipe}(u, v)\) such that \((u \rightarrow v) \in \text{Out}_u\). Initially these counters have value zero. For every \(v\) such that \((v \rightarrow u) \in \text{In}_v\), the Boolean variable \(\text{pending}_{\text{in}}(v)\) (initially set to \(\text{false}\)) indicates whether \(\text{pipe}(v, u)\) is one of the pipes in need of flushing for \(u\) to become active. Constants and variables carrying the subscript \(u\) in their names may be thought of as being part of task \(u\)'s "activation record", and do as such migrate along with \(u\) whenever it migrates.

Algorithm A_FIFO, given next for node \(n\), is an asynchronous algorithm to ensure the FIFO order of message delivery among migrating tasks. When listing the variables for this algorithm, only those carrying the subscript \(i\) are presented. The others, which refer to tasks, are omitted from the description. This same practice of only listing variables that refer to \(G\) is adopted everywhere in the book.
Algorithm A_FIFO:

Variables:
\[ \text{pipe}(u, v) = n, A(v) \text{ for all } (u \to v) \in D_r \text{ such that } A(u) = n; \]
\[ A(v) \text{ for all } v \in N_r. \]

Listing 2.3

Input:
\[ \text{msg}_i = \text{nil}. \]

Action when \( \text{active}_u \) and a decision is made to migrate \( u \) to \( n_j \):
\[ \text{active}_u := \text{false}; \]
for all \( (u \to v) \in \text{Out}_u \) do
\begin{itemize}
  \item Send \( \text{flush}(u, v, n_j) \) to \( A(v) \);
  \item pending_out := pending_out + 1
\end{itemize}
for all \( (v \to u) \in \text{In}_u \) do
\begin{itemize}
  \item Send \( \text{flush}_\text{request}(v,u) \) to \( A(v) \);
  \item pending_in := pending_in + 1;
  \item pending_in(v) := true
\end{itemize}
\[ A(u) := n_j; \]
Send \( u \) to \( n_j \).

Listing 2.4

Input:
\[ \text{msg}_i = u. \]

Action:
\[ A(u) := n_i. \]

Listing 2.5

Input:
\[ \text{msg}_i = \text{flush}(v, u, n_j). \]

Action:
\[ \text{if } A(u) = n \text{ then} \]
\[ \text{begin} \]
\[ A(v) := n_j; \]
\[ \text{Send } \text{flushed}(v,u,n) \text{ to } n_j; \]
\[ \text{if } \text{pending_in}_u(v) \text{ then} \]
\[ \text{begin} \]
\[ \text{pending_in}_u(v) := \text{false}; \]
\[ \text{pending_in}_u := \text{pending_in}_u - 1; \]
\[ \text{active}_u \]
\[ \begin{array}{l}
\quad \text{end} \\
\quad \text{else} \\
\quad \text{Send } flush(v,u,n) \text{ to } A(u).
\end{array} \]

**Listing 2.6**

Input:
\[ msg_i = \text{flush\_request}(u,v). \]

Action:
\[ \text{if } A_i(u) = n_i \text{ then} \]
\[ \begin{array}{l}
\quad \text{active}_u := \text{false}; \\
\quad \text{Send } flush(u,v,n_i) \text{ to } A(v); \\
\quad \text{pending\_out}_u := \text{pending\_out}_u + 1
\end{array} \]

**Listing 2.7**

Input:
\[ msg_i = \text{flushed}(u, v, n_i). \]

Action when \( A_i(u) = n_i \):
\[ \begin{array}{l}
\quad A_i(v) := n_i; \\
\quad \text{pipe}(u,v) := n_i, n_i; \\
\quad \text{pending\_out}_u := \text{pending\_out}_u - 1; \\
\quad \text{active}_u := (\text{pending\_in}_u = 0) \text{ and } (\text{pending\_out}_u = 0).
\end{array} \]

Algorithm \( A\_\text{FIFO} \) expresses, following the conventions established with Algorithm \( A\_\text{Template} \), the procedure described informally in Section 1.6.2. One important observation about Algorithm \( A\_\text{FIFO} \) is that the set \( N_0 \) of potential spontaneous senders of messages now comprises the nodes that concurrently decide to send active tasks to run elsewhere (cf. (2.3)), in the sense described in Section 1.6.2, and may then be such that \( N_0 = N \). In fact, the way to regard spontaneous initiations in Algorithm \( A\_\text{FIFO} \) is to view every maximal set of nodes concurrently executing (2.3) as an \( N_0 \) set for a new execution of the algorithm, provided every such execution operates on data structures and variables that persist (i.e., are not re-initialized) from one execution to another.

For completeness, next we give some of Algorithm \( A\_\text{FIFO} \)'s properties related to its correctness and performance.

**Theorem 2.1.**

For any two tasks \( u \) and \( v \) such that \( (u \rightarrow v) \in \text{Out}_n \), messages sent by \( u \) to \( v \) are delivered in the FIFO order.

**Proof:** Consider any scenario in which both \( u \) and \( v \) are active, and in this scenario let \( n_i \) be the node on which \( u \) runs and \( n_j \) the node on which \( v \) runs.
There are three cases to be analyzed in connection with the possible migrations of \( u \) and \( v \) out of \( n_i \) and \( n_j \), respectively.

In the first case, \( u \) migrates to another node, say \( n_j \), while \( v \) does not concurrently migrate, that is, the flush\((u,v,n_i)\) sent by \( n_i \) in \((2.3)\) arrives at \( n_j \) when \( A_j(v) = n_j \). A flushed\((u,v,n_j)\) is then by \((2.5)\) sent to \( n_i \), and may upon receipt cause \( u \) to become active if it is no longer involved in the flushing of any pipe (\( pending_in = 0 \) and \( pending_out = 0 \)), by \((2.7)\). Also, pipe\((u,v)\) is in \((2.7)\) set to \( n_i,n_j \), and it is on this pipe that \( u \) will send all further messages to \( v \) once it becomes active. These messages will reach \( v \) later than all the messages sent previously to it by \( u \) when \( u \) still ran on \( n_i \), as by \( G_j \)'s FIFO property all these messages reached \( n_j \) and were added to queue\(_i\) before \( n_j \) received the flush\((u,v,n_i)\).

In the second case, it is \( v \) that migrates to another node, say \( n_i \), while \( u \) does not concurrently migrate, meaning that the flush\_request\((u,v)\) sent by \( n_i \) to \( n_j \) in \((2.3)\) arrives when \( A_i(u) = n_j \). What happens then is that, by \((2.6)\), as pending\_out\(_i\) is incremented and \( u \) becomes inactive (if already it was not, as pending\_out\(_i\) might already be positive), a flush\((u,v,n_i)\) is sent to \( n_j \) and, finding \( A_j(v) \neq n_i \), by \((2.5)\) gets forwarded by \( n_i \) to \( n_j \). Upon receipt of this message at \( n_j \), a flushed\((u,v,n_j)\) is sent to \( n_i \), also by \((2.5)\). This is a chance for \( v \) to become active, so long as no further pipe flushings remain in course in which it is involved (\( pending_in = 0 \) and \( pending_out = 0 \) in \((2.5)\)). The arrival of that message at \( n_j \) causes pending\_out\(_i\) to be decremented in \((2.7)\), and possibly \( u \) to become active if it is not any longer involved in the flushing of any other pipe (\( pending_in = 0 \) and \( pending_out = 0 \)). In addition, pipe\((u,v)\) is updated to \( n_i,n_j \). Because \( u \) remained inactive during the flushing of pipe\((u,v)\), every message it sends to \( v \) at \( n_i \) when it becomes active will arrive at its destination later than all the messages it had sent previously to \( v \) at \( n_i \), as once again \( G_j \)'s FIFO property implies that all these messages must have reached \( n_j \) and been added to queue\(_i\) ahead of the flush\((u,v,n_i)\).

The third case corresponds to the situation in which both \( u \) and \( v \) migrate concurrently, say respectively from \( n_i \) to \( n_j \) and from \( n_j \) to \( n_i \). This concurrency implies that the flush\((u,v,n_i)\) sent in \((2.3)\) by \( n_i \) finds \( A_j(v) \neq n_i \) on its arrival (and is therefore forwarded to \( n_j \), by \((2.5)\)), and likewise the flush\_request\((u,v)\) sent in \((2.3)\) by \( n_j \) to \( n_i \) finds \( A_i(u) \neq n_i \) at its destination (which by \((2.6)\) does nothing, as the flush\((u,v,n_i)\) it would send as a consequence is already on its way to \( n_i \) or \( n_j \)). A flushed\((u,v,n_j)\) is sent by \( n_j \) to \( n_i \), where by \((2.7)\) it causes the contents of pipe\(_i\)(\(u,v\)) to be updated to \( n_i',n_j' \). The conditions for \( u \) and \( v \) to become active are entirely analogous to the ones we discussed under the previous two cases. When \( u \) does finally become active, any messages it sends to \( v \) will arrive later than the messages it sent previously to \( v \) when it ran on \( n_i \) and \( v \) on \( n_j \). This is so because, once again by \( G_j \)'s FIFO property, such messages must have reached \( n_j \) and been added to queue\(_i\) ahead of the flush\((u,v,n_i)\).

Let \(|\text{pipe}(u,v)|\) denote the number of nodes in pipe\((u,v)\). Before we state Lemma 2.2, which establishes a property of this quantity, it is important to note that the number of nodes in pipe\((u,v)\) is not to be mistaken for the number of nodes in \( n_i \)'s view of that pipe if \( n_i \) is the node on which \( u \) runs. This view, which we have denoted by pipe\((u,v)\), clearly contains at most two nodes at all times, by \((2.7)\). The former, on the other hand, does not have a precise meaning in the framework of any node considered individually, but rather should be taken in the context of a consistent global state (cf. Section 3.1).

**Lemma 2.2.**

For any two tasks \( u \) and \( v \) such that \( u \rightarrow v \) \( \in \text{Out}_v \), \(|\text{pipe}(u,v)| \leq 4 \) always holds.

**Proof:** It suffices to note that, if \( u \) runs on \( n_i \), \(|\text{pipe}(u,v)| \) is larger than the number of nodes in pipe\((u,v)\) by at most two nodes, which happens when both
u and v migrate concurrently, as neither of the two tasks is allowed to migrate again before the pipe between them is shortened. The lemma then follows easily from the fact that by (2.7) pipe(u,v) contains at most two nodes.

To finalize our discussion of Algorithm A_FIFO in this section, we present its complexity. This quantity, which we still have not introduced and will only describe at length in Section 3.2, yields, in the usual worst-case asymptotic sense, a distributed algorithm's "cost" in terms of the number of messages it employs and the time it requires for completion. The message complexity is expressed simply as the worst-case asymptotic number of messages that flow among neighbors during the computation ("worst case" here is the maximum over all variations in the structure of G, when applicable, and over all executions of the algorithm— cf. Section 3.2.1). The time-related measures of complexity are conceptually more complex, and an analysis of Algorithm A_FIFO in these terms is postponed until our thorough discussion of complexity measures in Section 3.2.

For a nonempty set \( K \subseteq N \) of tasks, we henceforth let \( m_K \) denote the number of directed edges in \( D_T \) of the form \((u \rightarrow v)\) or \((v \rightarrow u)\) for \( u \in K \) and \( v \in N \).

Clearly,

\[
m_K \leq \sum_{u \in K} (|In_u| + |Out_u|) \leq 2m_K.
\]

**Theorem 2.3.**

For the concurrent migration of a set \( K \) of tasks, Algorithm A_FIFO employs \( O(m_K) \) messages.

**Proof:** When a task \( u \in K \) migrates from node \( n_i \) to node \( n_j \), \( n_i \) sends \( |In_u| \) messages \( flush_request(v, u) \) for \((v \rightarrow u) \in In_u\) and \( |Out_u| \) messages \( flush(u,v,n_i) \) for \((u \rightarrow v) \in Out_u\). In addition, \( n_i \) receives \( |In_u| \) messages \( flush(v,u,n_i) \) for \((v \rightarrow u) \in In_u\) and some appropriate \( n_i \), and \( |Out_u| \) messages \( flushed(u,v,n_i) \) for \((u \rightarrow v) \in Out_u\) and some appropriate \( n_i \). Node \( n_i \) also sends \( |In_u| \) messages \( flushed(v,u,n_i) \) for \((v \rightarrow u) \in In_u\). Only \( flush \) messages traverse pipes, which by Lemma 2.2 contain no more than four nodes or three edges each. Because no other messages involving \( u \) are sent or received even if other tasks \( v \) such that \((v \rightarrow u) \in In_u\) or \((u \rightarrow v) \in Out_u\) are members of \( K \) as well, except for the receipt by \( n_i \) of one innocuous message \( flush_request(u, v) \) for each \( v \in K \) such that \((u \rightarrow v) \in Out_u\), the concurrent migration of the tasks in \( K \) accounts for \( O(m_K) \) messages.

The message complexity asserted by Theorem 2.3 refers to messages sent on the edges of \( G \), which is a complete graph. It would also be legitimate, in this context, to consider the number of interprocessor messages actually employed, that is, the number of messages that get sent on the edges of \( G_p \). In the case of fixed, deterministic routing (cf. Section 1.3), a message on \( G \) corresponds to no more than \( n - 1 \) messages on \( G_p \), so by Theorem 2.3 the number of interprocessor messages is \( O(nm_K) \). However, recalling our remark in Section 1.3 when we discussed the use of wormhole routing for flow control in multiprocessors, if the transport of interprocessor messages is efficient enough that \( G_p \) too can be regarded as a complete graph, then the message complexity given by Theorem 2.3 applies to interprocessor messages as well.

In addition to the asynchronous model we have been discussing so far in this section, another model related to \( G \)'s timing characteristics is the fully synchronous (or simply synchronous) model, for which the following two properties hold.

- All nodes are driven by a global time basis, referred to as the global clock, which generates time intervals (or simply intervals) of fixed, nonzero duration.
• The delay that a message suffers to be delivered between neighbors is nonzero and strictly less than the duration of an interval of the global clock.

The intervals generated by the global clock do not really need to be of the same duration, so long as the assumption on the delays that messages suffer to be delivered between neighbors takes as bound the minimum of the different durations.

The following is an outline of the functioning of a distributed algorithm, called a synchronous algorithm, designed under the assumptions of the synchronous model. The beginning of each interval of the global clock is indicated by a pulse. For \( s \geq 0 \), pulse \( s \) indicates the beginning of interval \( s \). At pulse \( s = 0 \), the nodes in \( N_0 \) send messages on some (or possibly none) of the edges directed away from them. At pulse \( s > 0 \), all the messages sent at pulse \( s - 1 \) have by assumption arrived, and then the nodes in \( N \) may compute and send messages out.

One assumption that we have tacitly made, but which should be very clearly spelled out, is that the computation carried out by nodes during an interval takes no time. Without this assumption, the duration of an interval would not be enough for both the local computations to be carried out and the messages to be delivered, because this delivery may take nearly as long as the entire duration of the interval to happen. Another equivalent way to approach this would have been to say that, for some \( d \geq 0 \) strictly less than the duration of an interval, local computation takes no more than \( d \) time, while messages take strictly less than the duration of an interval minus \( d \) to be delivered. What we have done has been to take \( d = 0 \). We return to issues related to these in Section 3.2.2.

The set \( N_0 \) of nodes that may send messages at pulse \( s = 0 \) has in the synchronous case the same interpretation as a set of potential spontaneous senders of messages it had in the asynchronous case. However, in the synchronous case it does make sense for nodes to compute without receiving any messages, because what drives them is the global clock, not the reception of messages. So a synchronous algorithm does not in principle require any messages at all, and nodes can still go on computing even if \( N_0 = \) \( \emptyset \). Nevertheless, in order for the overall computation to have any meaning other than the parallelization of \( n \) completely independent sequential computations, at least one message has to be sent by at least one node, and for a message that gets sent at the earliest pulse that has to take place at pulse \( s = d \) for some \( d \geq 0 \). What we have done has been once again to make the harmless assumption that \( d = 0 \), because whatever the nodes did prior to this pulse did not depend on the reception of messages and can therefore be regarded as having been done at this pulse as well. Then the set \( N_0 \) has at least the sender of that message as member.

Unrealistic though the synchronous model may seem, it may at times have great appeal in the design of distributed algorithms, not only because it frequently simplifies the design (cf. Section 4.3, for example), but also because there have been cases in which it led to asynchronous algorithms more efficient than the ones available (cf. Section 3.4). One of the chiefest advantages that comes from reasoning under the assumptions of the synchronous model is the following. If for some \( d > 0 \) a node \( n_i \) does not receive any message during interval \( s \) for some \( s \geq d \), then surely no message that might "causally affect" the behavior of \( n_i \) at pulse \( s + 1 \) was sent at pulses \( s - d, \ldots, s \) by any node whose shortest distance to \( n_i \) is at least \( d \). The "causally affect" will be made much clearer in Section 3.1 (and before that used freely a few times), but for the moment it suffices to understand that, in the synchronous model, nodes may gain information by just waiting, i.e., counting pulses. When designing synchronous algorithms, this simple
observation can be used for many purposes, including the detection of
termination in many cases (cf., for example, Sections 2.2.2 and 2.2.3).
It should also be clear that every asynchronous algorithm is also in essence a
synchronous algorithm. That is, if an algorithm is designed for the
asynchronous model and it works correctly under the assumptions of that
model, then it must also work correctly under the assumptions of the
synchronous model for an appropriate choice of interval duration (to
accommodate nodes' computations). This happens because the conditions
under which communication takes place in the synchronous model is only one
of the infinitely many possibilities that the asynchronous model allows. We
treat this issue in more detail in Section 3.3. The converse of this implication
(i.e., that synchronous algorithms run correctly in the asynchronous model)
can also be achieved with appropriate algorithm transformation, and is not at
all immediate as its counterpart. This transformation lends support to our
interest in the synchronous model and is our subject in Section 5.3, after we
return to it in Sections 3.3 and 3.4.

Our last topic in this section is the presentation of Algorithm S_Template,
which sets the conventions on how to describe a synchronous algorithm and is
used as a template throughout the book. The prefix S_, similarly to the
asynchronous case discussed earlier, indicates that the algorithm is
synchronous, and is used in all synchronous algorithms we present. For \( s \geq 0 \)
and \( n_i \in N \), in Algorithm S_Template MSG\(_i\)(s) is either the empty set (if \( s = 0 \))
or denotes the set of messages received by \( n_i \) during interval \( s - 1 \) (if \( s > 0 \)),
which may be empty as well. The algorithm for \( n_i \) is given next. As with
Algorithm A_Template, Algorithm S_Template too is given for the case in
which \( G \) is a directed graph. The undirected case is obtained by simply
replacing In and Out with Inc throughout the algorithm.

**Algorithm S_Template:**

**Variables:**

Variables used by \( n_i \), and their initial values, are listed here.

**Listing 2.8**

**Input:**

\[ s = 0, \text{MSG}(0) = \emptyset \]

**Action if** \( n_i \in N_0 \):

Do some computation;

Send one message on each edge of a (possibly empty) subset of
\( \text{Out}_i \).

**Listing 2.9**

**Input:**

\[ s > 0, \text{MSG}(1), \ldots, \text{MSG}(s) \] such that \( \text{origin}(msg) = c_k \in \text{In}_i \)

with \( 1 \leq k \leq |\text{In}_i| \) for

\[ msg \in \bigcup_{r=1}^{s} \text{MSG}(r). \]

**Action:**

Do some computation;
Send one message on each edge of a (possibly empty) subset of Out.

As in the case of Algorithm A_Template, Algorithm S_Template is presented as a set of input/action pairs whose actions are numbered for ease of reference (2.8 corresponds to $s = 0$ and 2.9 to $s > 0$). The inputs now include information from the global clock (in the form of the nonnegative integer $s$), which is, as we have seen, what really drives the nodes. The atomicity of the actions comes as a consequence of the characteristics of the synchronous model, because no node performs more than one action per interval of the global clock. In fact, it is simple to see that every node performs exactly one action per interval of the global clock, because actions are now unconditional, that is, in describing Algorithm S_Template we have done away with the Boolean conditions that Algorithm A_Template inherited from the guard’s of Algorithm Task_t of Section 1.4. The reason why we could do this is that such conditions are in the synchronous case evaluated only at the occurrence of pulses, and this can be treated inside the action itself (through the use of ifs, as opposed to the use of when’s in the asynchronous case).

Another important observation regarding Algorithm S_Template is that we allow $n_i$ to have access, during its computation at interval $s > 0$, to all the sets $MSG_i(1),...,MSG_i(s)$. Although normally only $MSG_i(s)$ is needed, the greater generality is useful for our purposes in various situations, as for example in Sections 2.2.3 and 3.3.

**2.2 Computations on anonymous systems**

The system represented by the graph $G$ is said to be an anonymous system when its nodes do not have identifications that they can use in their computations. Of course, we as outside observers can still make use of the identifications $n_1,...,n_n$ in describing the anonymous system, the computations that run on it, and the properties of those computations. The nodes themselves, however, cannot have access to such identifications for use in the algorithm, not even to identify a neighbor as the source or the destination of a message. In an anonymous system, all that is known to a node $n_i$ are the sets $In_i$ and $Out_i$ of edges (Inc, in the undirected case), so messages have to be received and sent over these edges without explicit mention to the nodes on the other side, whose identifications are unknown. When receiving a message $msg$, the only information related to the origin of $msg$ that $n_i$ can use is the identification of the edge on which the message arrived, and this is denoted by $\text{origin}(msg)$, as we discussed in Section 2.1. The reader should check that Algorithms A_Template and S_Template of Section 2.1 were written in this fashion, so they can be used directly to express algorithms on anonymous systems.

The study of computations on anonymous systems is interesting from at least two perspectives. First of all, this study provides an opportunity to investigate the limits of what can be computed distributedly when nodes do not have, and cannot possibly obtain, complete information on the overall structure of $G$. The second perspective is that of systems that really should be regarded as anonymous, as many systems represented by massively parallel models that in fact can be viewed as performing distributed computations (cf. Section 10.2 for examples).

One of the foremost consequences of assuming that a system is anonymous is that the algorithm describing the computation to be carried out by a node must be the same for all nodes. The reason why this property must hold is that differences in the algorithms performed by the nodes might provide a means to establish identifications that the nodes would then be able to use in their computations, in which case the system would no longer be anonymous.

Our discussion throughout Section 2.2, will be limited to the cases in which $G$ is an undirected graph with one single cycle, that is, an undirected ring. In the case of a ring, Inc,
has exactly two members for all $n \in N$, which we let be called \textit{left} and \textit{right}. If every edge $(n_i, n_j)$ is such that $(n_i, n_j) = \text{left} = \text{right}$, then we say that the ring is \textit{locally oriented}, or, equivalently, that the assignment of denominations to edges locally at the nodes establishes a \textit{local orientation} on the ring. Equivalently, this can be expressed by rephrasing the condition as $(n_i, n_j) = \text{left} = \text{right}$ for all $(n_i, n_j) \in E$.

Section 2.2.1 contains a discussion of two impossibility results under the assumption of anonymity. These two results refer to computations of Boolean functions and to the establishment of local orientations under certain assumptions on $n$, the number of nodes in the ring. The remaining two sections contain algorithms to compute Boolean functions (Sections 2.2.2) and to find a local orientation (Sections 2.2.2 and 2.2.3) when the conditions leading to the impossibility results of Section 2.2.1 do not hold.

2.2.1 Some impossibility results

Let $f$ be a Boolean function of the form

$$f : \{\text{true, false}\}^* \to \{\text{true, false}\}.$$

In this section, we consider algorithms to compute $f$ when the $n$ Booleans that constitute its arguments are initially scattered throughout the nodes, one per node, in such a way that at the end of the algorithm every node has the same value for $f$ (we say that such an algorithm \textit{computes $f$ at all nodes}). Naturally, the assignment of arguments to nodes has to be assumed to be given initially, because an anonymous system cannot possibly perform such an assignment by itself.

The first impossibility result that we discuss is given by Theorem 2.4, and is related to the availability of $n$ to be used by the nodes in their computations.

\textbf{Theorem 2.4.} 

\textit{No synchronous algorithm exists to compute $f$ at all nodes if $n$ is not known to the nodes.}

\textbf{Proof:} We show that any synchronous algorithm that computes $f$ in the absence of information on $n$ must in some cases fail, that is, we show that such an algorithm does not necessarily compute $f$ at all nodes.

For consider an algorithm to compute $f$ when $n$ is not known to the nodes. This algorithm must function independently of $n$, therefore for rings with all numbers of

\begin{figure}
\centering
\includegraphics[width=\textwidth]{figure21.png}
\caption{This is the $2v \cdot (2T/3 + 1)$-node ring used in the proof of Theorem 2.4, here shown for $v = 3$ and $T = 3$. Each of the three portions in the upper half comprising three contiguous nodes each is assigned $f$'s arguments according to $a_t$. Similar portions in the lower half of the ring follow assignment $a_t$.}
\end{figure}

nodes. In particular, for a ring with $n = v \geq 3$ nodes, let $a_r$ and $a_t$ be assignments of $f$'s arguments to nodes, i.e.,

$$a_r, a_t : \{n_1, \ldots, n_v\} \to \{\text{false, true}\},$$

such that
Furthermore, let $T_f$ and $T_t$ be the numbers of pulses that the algorithm spends in computing $f$ for, respectively, assignments $a_f$ and $a_t$. Let $T$ be such that $T \geq \max\{T_f, T_t\}$.

The next step is to consider a ring with $n = 2v(2l/3 + 1)$ nodes, for which the algorithm must also work, and to assign arguments to the nodes as follows. Divide the ring into two connected halves, and within each half identify $2l/3 + 1$ portions, each with $v$ contiguous nodes. To each such portion in one of the halves assign arguments as given by $a_f$. Then use $a_t$ to do the assignments to each of the portions in the other half (Figure 2.1).

Because the number of portions in each half is odd, we can identify a middle portion in each of the halves. Also, except for the nodes at either end of the two halves, every node is in the larger ring connected as it was in the smaller one (i.e., the Booleans assigned to a node's neighbors are the same in the two rings). In the synchronous model, it takes at least $d$ pulses for a node to causally affect another that is $d$ edges apart on a shortest path, so nodes in the middle portions of both halves cannot be causally affected by any other node in the other half within $T$ pulses of the beginning of the computation. What these considerations imply is that the nodes in the middle portion of the half related to $a_f$ will by pulse $T$ have terminated and proclaimed the value of $f$ to be false, because this is what happened by assumption under the same circumstances on the smaller ring. Similarly, nodes in the middle portion of the half related to $a_t$ will have terminated and proclaimed $f$ to have value true within $T$ pulses of the beginning of the computation.

Corollary 2.5.

*No algorithm exists to compute $f$ at all nodes if $n$ is not known to the nodes.*

**Proof:** This is a direct consequence of our discussion in Section 2.1, where we mentioned that every asynchronous algorithm easily yields an equivalent synchronous algorithm. So, if an asynchronous algorithm existed to compute $f$ at all nodes in the absence of information on $n$ at the nodes, then the resulting synchronous algorithm would contradict Theorem 2.4. If $n$ is known to the nodes, then $f$ can be computed at all nodes by a variety of algorithms, as we discuss in Section 2.2.2.

The second impossibility result that we discuss in this section is related to establishing a local orientation on the ring when, for $n_i \in N$, the identifications $left_i$ and $right_i$ are not guaranteed to yield a local orientation initially. This problem is related to the problem of computing $f$ we discussed previously in the following manner. At node $n_i$, the positioning of $left_i$ and $right_i$ with respect to how its neighbors' edge identifications are positioned can be regarded as constituting a Boolean input. Establishing a local orientation for the ring can then be regarded as computing a function $f$ on these inputs and then switching the denominations of the two edges incident to $n_i$ if the value it computes for $f$ turns out to be, say, false. Now, this function is not in general expected to yield the same value at all nodes, and then Corollary 2.5 would not in principle apply to it. However, another Boolean function, call it $f'$, can be computed easily once $f$ has been computed. This function has value true if and only if the ring is locally oriented, and this is the value it would be assigned at each node right after that node had computed $f$ and chosen either to perform the switch in edge identifications or not to. Clearly, $f'$ is expected to be assigned the same value at all nodes, and then by Corollary 2.5 there is no algorithm to compute it at all nodes in the absence of information on $n$. As a consequence, there is no algorithm to compute $f$ either.

Even when $n$ is known to the nodes, there are cases in which no algorithm can be found to establish a local orientation on the ring. Theorem 2.6 gives the conditions under which this happens.
Theorem 2.6.
No synchronous algorithm exists to establish a local orientation on the ring if \( n \) is even.

**Proof:** Our argument is to show that any synchronous algorithm to establish a local orientation on the ring fails in some cases if \( n \) is even. To do so, we let \( n = 2v \) for some \( v \geq 2 \), and then consider the following arrangement of left and right for all \( n_i \in N \). For \( 1 \leq i \leq v - 1 \), we let
\[
right_i = left_{i+1},
\]
and for \( v + 2 \leq i \leq 2v \) we let
\[
right_i = left_{i-1}.
\]

Clearly, this arrangement also implies
\[
left_i = left_{2v},
\]
and
\[
right_i = right_{2v+1},
\]
so the ring is not locally oriented.

Now we consider a mapping \( \varphi : \{1, \ldots, 2v\} \rightarrow \{1, \ldots, 2v\} \)
\[
\text{such that, for } 1 \leq i \leq 2v,
\]
\[
\varphi(i) = 2v+1-i\]
for which it clearly holds that (Figure 2.2). This mapping is also such that, for \( 1 \leq i \leq 2v \), if \( n_i \) sends a message at a certain pulse (or receives a message during the corresponding interval) on edge left, or edge right, then \( n_i \) does exactly the same at the same pulse, respectively on edge left or edge right.

Consequently, \( n_i \) and \( n_{\varphi(i)} \) reach the same conclusion on whether they should switch their incident edges' identifications or not, and the ring continues to be not locally oriented.

**Corollary 2.7.**
No algorithm exists to establish a local orientation on the ring if \( n \) is even.

**Proof:** The proof here is entirely analogous to that of Corollary 2.5. If \( n \) is known to the nodes and is odd, then a local orientation can be established on the ring. We give algorithms to do this in Sections 2.2.2 and 2.2.3.
2.2.2 Boolean-function computations

When $n$ is known to the nodes, Corollary 2.5 does not apply and the function $f$ introduced in Section 2.2.1 can be computed at all nodes. Also, if such a function is computed with the aim of eventually establishing a local orientation on the ring, then $n$ has to be odd for Corollary 2.7 not to apply.

In this section, we start by presenting Algorithm $A_{\text{Compute } f}$, which is an asynchronous algorithm to compute $f$ at all nodes when $n$ is known to the nodes. In addition, we present this algorithm in such a way that, if $n$ is odd, then it may be used almost readily to establish a local orientation on the ring as well.

In Algorithm $A_{\text{Compute } f}$, $b \in \{\text{false}, \text{true}\}$ denotes $f$'s argument corresponding to $n_i \in N$.

In order for the algorithm to be also suitable to the determination of a local orientation on the ring, the messages that it employs carry the pair of Booleans comprising one argument of $f$ and a Boolean constant.

So far as computing $f$ goes, the essence of Algorithm $A_{\text{Compute } f}$ is very simple. If $n_i \in N_0$, or upon receiving the first message if $n_i \notin N_0$, $n_i$ sends the pair $(b_i, \text{false})$ on $\text{left}_i$ and the pair $(b_i, \text{true})$ on $\text{right}_i$. For each of the $\lceil n/2 \rceil$ messages it receives on each of the edges incident to it, $n_i$ records the Booleans contained in the message and sends them onward on the edges opposite to those on which they were received. After all these messages have been received, $n_i$ has the Booleans originally assigned to every node and may then compute $f$ locally.

Node $n_i$ employs two variables to count the numbers of messages received, respectively $\text{count-left}_i$ and $\text{count-right}_i$ for $\text{left}_i$ and $\text{right}_i$. Initially, these counters have value zero. In addition, $n_i$ employs the $n$ Boolean variables $b_i^j$ to record the values of $b_i, \ldots, b_n$ when they are received in messages (if $j \neq i$) for $1 \leq j \leq n$. Initially, $b_1^j = b_i$ (the others do not need any initial value to be set). Another variable $j_i$ is used to contain the subscripts to these variables.

Because Algorithm $A_{\text{Compute } f}$ has to be exactly the same for all nodes in $N$, another Boolean variable, $\text{initiated}_i$, (initially set to false), is employed by $n_i$ to indicate whether $n_i \in N_0$ or not. This variable is set to true when $n_i$ starts its computation if it is a member of $N_0$. Nonmembers of $N_0$ will have this variable equal to false upon receiving the first messages, and will then know that first of all it must send messages out. In the absence of anonymity, sometimes it is simpler to specify an algorithm for $n_i \in N_0$ and another for $n_i \notin N_0$.

Algorithm $A_{\text{Compute } f}$:

**Variables:**
- $\text{count-left}_i = 0$
- $\text{count-right}_i = 0$
- $b_i^k = b_i$ (if $k = 1$) for $1 \leq k \leq n$
- $j_i = 1$
- $\text{initiated}_i = \text{false}$

Listing 2.10

**Input:**
- $msg_i = \text{nil}$.

**Action if $n_i \in N_0$**
- $\text{initiated}_i := \text{true}$;

Send $(b_i^1, \text{false})$ on $\text{left}_i$;
Listing 2.11

Input:
$\text{msg}_i = (b,B)$

Action:

if not $\text{initiated}$, then

begin

$\text{initiated}_i := \text{true}$;

Send $b^1_\frac{1}{2}$, $\text{false}$ on left;

Send $b^1_\frac{1}{2}$, $\text{true}$ on right;

end;

if $\text{origin}_i(\text{msg}_i) = \text{left}$, then

begin

$count_{\text{left}}_i := count_{\text{left}}_i + 1$;

$j_i := j_i + 1$;

if $j_i \leq n$ then

$b_{ji} := b$;

if $count_{\text{left}}_i \leq \lfloor n/2 \rfloor - 1$ then

Send $\text{msg}_i$ on right;

end;

if $\text{origin}_i(\text{msg}_i) = \text{right}$, then

begin

$count_{\text{right}}_i := count_{\text{right}}_i + 1$;

$j_i := j_i + 1$;

if $j_i \leq n$ then

$b_{ji} := b$;
An instructive observation at this point is that Algorithm \texttt{A\_Compute\_f} is indeed an algorithm for anonymous rings. Nowhere in the algorithm are the identities of the nodes mentioned, except in the description of (2.10), but this is only for notational consistency with Algorithm \texttt{A\_Template}, because in any event the set \( N_i \) is determined by an "external agent." In fact, it is because of the system's anonymity that the \( B \)'s that \( n \) receives have to be placed by (2.11) in the variables \( b_1^2, \ldots, b_n^2 \) irrespective of their original senders, which would be simpler if the denominations of those senders could be used by the algorithm. In addition, the algorithm does make use of \( n \), as anticipated by Corollary 2.5.

Let us now examine Algorithm \texttt{A\_Compute\_f} carefully. During \( n \)'s computation, it receives the messages originally sent to it by its neighbors by (2.10) or (2.11), and whatever those neighbors forward to it by (2.11). Because by (2.11) a node only forwards to each of its neighbors \( \lceil n/2 \rceil - 1 \) messages, \( n \) actually receives \( 2 \lceil n/2 \rceil \) messages, of which the last two it does not forward. Upon receipt of the last of the \( 2 \lceil n/2 \rceil \) messages, \( n \) has either \( n \) (if \( n \) is odd) or \( n + 1 \) (if \( n \) is even) arguments of \( f \), which it may then compute. Although it may be possible to modify the algorithm a little bit to ensure that exactly \( n \) arguments are received if \( n \) is even as well (cf. Exercise 1), as presented the last argument received is a repetition and may be dropped (as in (2.11)).

In many cases, it may only be possible to compute \( f \) if the information \( n \) receives is organized more orderly than as in Algorithm \texttt{A\_Compute\_f}. In other words, unless \( f \) is invariant with respect to the order of its arguments (as in the case of the \texttt{AND} and \texttt{OR} functions, for example), then the variables \( b_1^2, \ldots, b_n^2 \) have to be replaced with two sets of similar variables, each with \( \lceil n/2 \rceil \) variables to accommodate the Booleans received from each of \( n \)'s neighbors. In addition, if such an invariance does not hold, then the edges in \( E \) have to be assumed to be FIFO. Even so, however, because the system is anonymous \( f \) can only be computed if it is invariant under rotations of its arguments.

As we mentioned earlier, Algorithm \texttt{A\_Compute\_f} can also be used to provide the ring with a local orientation, and this is the role of the \( B \)'s that get sent along with every message. When the algorithm is used with this purpose, then the \( b \)'s have no role and the \( B \)'s are treated as follows at the step in which \( f \) would be computed in (2.11). A \( B \) that \( n \) receives indicates either that its original sender had its \textit{left} and \textit{right} edges positioned like \textit{left} and \textit{right}; (if \( B = \texttt{true} \) is received on \textit{left} or \( B = \texttt{false} \) is received on \textit{right} or positioned otherwise (if \( B = \texttt{false} \) is received on \textit{left} or \( B = \texttt{true} \) is received on \textit{right}). In either case, so long as \( n \) is odd (and \( n \) has to be odd, by Corollary 2.7), \( n \) can decide whether its-edges are positioned like those of the majority of the nodes, in which case it maintains their positioning, or not, in which case it reverses their positioning. The result of these decisions system-wide is clearly to establish a local orientation on the ring. (Note that in this case Algorithm \texttt{A\_Compute\_f} would have to be modified to treat the \( B \)'s, not the \( b \)'s, \( n \) receives—cf. Exercise 3.)

Because each node receives \( 2 \lceil n/2 \rceil \) messages during the computation, the total number of messages employed by the algorithm is \( 2n \lceil n/2 \rceil \), and its message complexity is clearly \( O(n^2) \).
In Section 3.2.1, we return to Algorithm A_Compute_f to discuss its time-related complexity measures. In the remainder of this section and in Section 2.2.3, we show that synchronous algorithms exist whose message complexities are significantly lower than that of Algorithm A_Compute_f, so long as the generality of this algorithm can be given up. The synchronous algorithm that we discuss next is specific to computing the AND function, while the one we discuss in Section 2.2.3 is specific to providing the ring with a local orientation. The key ingredient in obtaining the more efficient synchronous algorithm is that the AND function can be assumed to be true unless any of its arguments are false. In the synchronous case, this observation can be coupled with the assumptions of the synchronous model as follows. Only nodes with false arguments send their argument to neighbors. The others simply wait to receive a false or long enough to know that any existing false would already have reached them. In either case, computing the AND is a simple matter. Algorithm S_Compute_AND embodies this strategy and is given next. In this algorithm, $N_0 = N$ and a Boolean variable $f_i$ (initially set to true) is employed by $n_i$ to store the result of evaluating the AND function.

**Algorithm S_Compute_AND**

**Variables:**

$f_i = \text{true}$.  

**Listing 2.12**

**Input:**

$s = 0$, $MSG(0) = \emptyset$

**Action if $n_i \in N_0$:**

if $b_i = \text{false}$ then

begin

$f_i := \text{false}$:

Send $b_i$ on left and on right.

end.

**Listing 2.13**

**Input:**

$0 < s < \lfloor n/2 \rfloor$, $MSG(s)$.

**Action:**

if $f_i$ then

if $MSG(s) \neq \emptyset$ then

begin

$f_i := \text{false}$;

if there exists $msg \in MSG(s)$ such that
The synchronous algorithm proceeds in iterations, each one comprising two phases. Initially, all nodes are said to be active, and the goal of each of the iterations is to reduce the number of active nodes. During an iteration, nodes that are not active function solely as message relays, so that the computation can always be looked at as being carried out on a ring containing the active nodes only, called the active ring. Iterations proceed until exactly one active node remains or until an active ring is reached which is locally oriented (in this case with more than one active node). A local orientation can then be established on the entire ring by the last active node.

The number of iterations that the algorithm requires depends largely on how active nodes are eliminated from one iteration to the next. In Algorithm S_Locally_Orient, given next, this elimination takes place as follows. In the first phase of an iteration, segment ends are identified and selected to be the only active nodes to remain through to the next iteration. By our preceding discussion, the number of center nodes must be odd, and then so must the number of nodes in every active ring, thereby guaranteeing the feasibility of every iteration. The last iteration is characterized by the absence of segment ends among the active nodes. Because at each iteration segments with an even number of nodes do not contribute with any active node to the next iteration, and considering that segments with odd numbers of nodes have at least three nodes each, clearly the number of iterations required is no larger than \( \left\lfloor \log_2 n \right\rfloor = O(\log n) \).

Letting \( \sigma \geq O \) indicate the pulses within each of the iterations, the following is how the two aforementioned phases within an iteration are implemented. At pulse \( \sigma = O \), node \( n_i \), if active, sends \textit{token} on edge \textit{right}. Active nodes then idle throughout the following \( n - 1 \) pulses.

\[
\text{origin}(msg) = \text{left}, \text{ then }
\text{Send msg on right;}
\]

\[
\text{if there exists msg} \in \text{MSG(s)} \text{ such that }
\text{origin}(msg) = \text{right, then }
\text{Send msg on left;}
\]

end.

If \( n_i \) is not such that \( b_i = \text{false} \), then the largest number of pulses that can go by before \( n_i \) concludes that \( f_i \), cannot be changed from its initial value of \textit{true} is \(|n/2|\), so that after pulse \( s = \lfloor n/2 \rfloor = O(n) \) no further computation has to be performed and the algorithm may terminate. By (2.12) and (2.13), \( n \) sends at most two messages during its computation, either initially if \( b_i = \text{false} \), by (2.12), or upon receiving the first message, if any messages are at all received, by (2.13). Clearly, then, the message complexity of Algorithm S_Compute_AND, is \( O(n) \).

2.2.3 Another algorithm for local orientation
In addition to Algorithm S_Compute_AND, another example of how to employ many fewer messages than those required by Algorithm A_Compute_f comes from considering a synchronous algorithm tailored specifically to establishing a local orientation on the ring. By Theorem 2.6, such an algorithm may only exist if \( n \) is odd, as we assume henceforth in this section.

The basic strategy behind this algorithm employs the following terminology. Say that two nodes \( n_i \) and \( n_j \) are segment ends if \((n_i, n_j) \in E \) and furthermore \((n_i, n_j) = \text{left}, \text{ left} \). Segment ends delimit segments, which are subsets of \( N \) inducing connected subgraphs of \( G \) with at least two nodes. If \( n_i \) and \( n_j \) are segment ends and \((n_i, n_j) \in E \), then \( n_i \) and \( n_j \) belong to different segments, unless the number of segments in the ring is exactly one. Clearly, a locally oriented ring contains no segment ends, while a ring that is not locally oriented contains a nonzero even number of segment ends, and half as many segments. Because \( n \) is odd, an odd number of segments must have an odd number of nodes each.

The synchronous algorithm proceeds in iterations, each one comprising two phases. Initially, all nodes are said to be active, and the goal of each of the iterations is to reduce the number of active nodes. During an iteration, nodes that are not active function solely as message relays, so that the computation can always be looked at as being carried out on a ring containing the active nodes only, called the active ring. Iterations proceed until exactly one active node remains or until an active ring is reached which is locally oriented (in this case with more than one active node). A local orientation can then be established on the entire ring by the last active node.

The number of iterations that the algorithm requires depends largely on how active nodes are eliminated from one iteration to the next. In Algorithm S_Locally_Orient, given next, this elimination takes place as follows. In the first phase of an iteration, segment ends are identified on the active ring. Then, in the second phase, the nodes, called center nodes, occupying the central positions in the segments having an odd number of nodes are identified and selected to be the only active nodes to remain through to the next iteration. By our preceding discussion, the number of center nodes must be odd, and then so must the number of nodes in every active ring, thereby guaranteeing the feasibility of every iteration. The last iteration is characterized by the absence of segment ends among the active nodes. Because at each iteration segments with an even number of nodes do not contribute with any active node to the next iteration, and considering that segments with odd numbers of nodes have at least three nodes each, clearly the number of iterations required is no larger than \(|\log_3 n| = O(\log n)\).

Letting \( \sigma \geq O \) indicate the pulses within each of the iterations, the following is how the two aforementioned phases within an iteration are implemented. At pulse \( \sigma = O \), node \( n_i \), if active, sends \textit{token} on edge \textit{right}. Active nodes then idle throughout the following \( n - 1 \) pulses,
while nodes that are not active simply relay token onward if they at all receive it. An active node $n_i$ that by pulse $\sigma = n$ has not received token on left, is a segment end, and at pulse $\sigma = n+1$ through pulse $\sigma = 2n - 1$, active nodes forward the integer $z + 1$ upon receiving integer $z$, for some $z \geq 0$, while the other nodes continue to function as relays. An active node $n_i$ that by pulse $\sigma = 2n$ has received the same integer over left and right, during the same interval is a center node. This is the last iteration if $n_i$ did not receive any message during intervals $n$ through $2n - 1$, otherwise only center nodes remain active for the next iteration. A message orient is sent on, say, left, by an active node $n_i$ after the last iteration. This message, if received on left by a node $n_i$ that is not active, causes left and right to be interchanged.

The reader should notice that the characteristics of the synchronous model are used profusely in this strategy to establish a local orientation. Indeed, both the determination of segment ends and of center nodes rely heavily on the assumed synchronism, as does the determination of when an iteration is the last one. In Algorithm $S_{\text{Locally Orient}}$, $k$ identifies the iteration and is then such that $1 \leq k \leq K$, where $K$ is the last iteration, therefore such that $K \leq \lceil \log_3 n \rceil$. Pulses within the $k$th iteration are numbered $s = 2n(k - 1) + \sigma$, that is, from $s = 2n(k - 1)$ through $s = 2nk$. After the last iteration, additional $n - 1$ pulses must elapse before termination. The only variable employed by $n_i$ is the Boolean variable active, initially set to true, used to indicate whether $n_i$ is active. Because initially active = true for all $n_i \in N$, in this algorithm $N_0 = N$. Because $K$ has to be determined as the algorithm progresses, it is assumed to be equal to infinity initially.

**Algorithm $S_{\text{Locally Orient}}$:**

**Variables:**

active $= \text{true}$.

**Listing 2.14**

**Input:**

$s = 2n(k - 1), \text{MSG}(s) =$ .

**Action (if } n_i \in N_0, \text{ for } k = 1):$

if active, then

Send token on right.

**Listing 2.15**

**Input:**

$2n(k - 1) + 1 \leq s \leq 2nk - n - 1, \text{MSG}(s)$.

**Action:**

if not active, then

begin

if there exists token $\in \text{MSG}(s)$ such that origin(token) = left, then

Send token on right;

if there exists token $\in \text{MSG}(s)$ such that origin(token) = right, then


Send token on left.

Listing 2.16

Input:
\[ s = 2nk - n, \text{MSG}(2n(k - 1) + 1), \ldots, \text{MSG}(2nk - n). \]

Action:
if active, then

- if there does not exist token \( MSG(r) \) such that \( \text{origin}(token) = \text{left} \), then
  - Send 0 on right.

Listing 2.17

Input:
\[ 2nk - n + 1 \leq s \leq 2nk - 1, \text{MSG}(s).s \]

Action:
if active, then

- begin
  - if there exists \( z \in \text{MSG}(s) \) such that \( \text{origin}(z) = \text{left} \), then
    - Send \( z + 1 \) on right;
  - if there exists \( z \in \text{MSG}(s) \) such that \( \text{origin}(z) = \text{right} \), then
    - Send \( z + 1 \) on left;
- end

else

- begin
  - if there exists \( z \in \text{MSG}(s) \) such that \( \text{origin}(z) = \text{left} \), then
    - Send \( z \) on right;
  - if there exists \( z \in \text{MSG}(s) \) such that \( \text{origin}(z) = \text{right} \), then
    - Send \( z \) on left;
- end.

Listing 2.18

Input:
\[ s = 2nk, \text{MSG}(2nk - n + 1), \ldots, \text{MSG}(2nk). \]
Action:
if active, then
  if \( \text{MSG}(r) = \emptyset \) for all \( r \in \{2nk - n, \ldots, 2nk - 1\} \) then
    begin
      \( K := k; \)
      Send orient on left;
    end
  else
    if there do not exist \( r \in \{2nk - n, \ldots, 2nk - 1\} \) and \( z_1, z_2 \in \text{MSG}(r) \) with \( z_1 = z_2 \) such that \( \text{origin}(z_1) = \text{left} \), and \( \text{origin}(z_2) = \text{right} \), then
      active, := false.

Listing 2.19

Input:
\( 2nK + 1 \leq s \leq 2nK + n - 1, \text{MSG}(s) \)
Action:
if not active, then
  begin
    if there exists orient \( \in \text{MSG}(s) \) such that \( \text{origin}(\text{orient}) = \text{left} \), then
      Interchange left, and right;
    Send orient on left;
  end.

In Algorithm \( S_{\text{Locally Orient}} \), (2.14) implements the sending of token at the beginning of each iteration, while in (2.15) the relaying of token by nodes that are not active appears. In (2.16), segment ends are identified and initiate the propagation of integers, which are relayed as appropriate by (2.17). Center nodes are identified in (2.18), which, in the last iteration, also includes the propagation of orient, relayed onward by (2.19). In no action does a node send more than two messages, and then the number of messages per iteration is clearly \( O(n) \). It follows from our earlier determination of the maximum number of iterations that the message complexity of Algorithm \( S_{\text{Locally Orient}} \) is \( O(n \log n) \).

2.3 The role of knowledge in distributed computations

The notion of knowledge is a notion of many possible meanings, but even the simplest algorithms we have seen so far in this chapter indicate that much of what distributed computations do is, in some sense, to collectively manipulate the system's knowledge so that at the end of the computation what nodes "know" individually relates in some way to the computation's original goal. During the past decade, various interesting hints at how notions related to knowledge might be used in the design and analysis of distributed algorithms were envisaged. Although today the interest in such an approach has waned somewhat, a few interesting insights were obtained that can be expressed in a particularly simple fashion when viewed from the standpoint of knowledge in the system.
Our goal in this section is to finalize the chapter by presenting some of these insights, which will be referred back to in forthcoming chapters for the sake of illustration within the context of those chapters. The ideal approach to our discussion in this section would be that of a logician, but naturally we refrain from doing that, especially because the necessary background to undertake such an approach intersects what is expected of a reader of this book very narrowly. Rather, we approach the subject quite informally, aiming essentially at conveying some of its intuitive underpinnings.

If $\mathcal{P}$ denotes a sentence (in the logical sense), then we denote the notion that a node $n_i$ knows $\mathcal{P}$ by $K_{i}\mathcal{P}$, where, loosely, $K_i$ is an operator indicating knowledge by $n_i$. Normally, only true sentences are assumed to be knowable, that is, in order for $\mathcal{P}$ to be known by $n_i$ it is necessary that $\mathcal{P}$ be true, giving rise to the axiom

$$K_i\mathcal{P} \rightarrow \mathcal{P}.$$  

(If $A$ and $B$ are two formulas, in the usual logical sense, then $A \rightarrow B$ is equivalent to $\neg A \lor B$.) Every distributed algorithm embodies various steps whereby the knowledge status of the nodes evolves. For example, if $n_i$ sends a message containing a true sentence $\mathcal{P}$ to $n_j \in \text{Neig}_i$, then $K_j\mathcal{P}$ holds as early as when the message is sent, but $K_j\mathcal{P}$ may happen to hold only from the time of receipt of the message onward, and then $K_jK_i\mathcal{P}$ also holds.

Despite the simplicity of such a notion of knowledge by $n_i$, it contains not too evident idiosyncrasies that include limits on what $n_i$ may know. This has been illustrated in the literature in the following anecdotic fashion.

"In a class with daily meetings the teacher announces, by the end of a Friday class, that there will be an unexpected exam in the following week. The students reason over the possibilities during the weekend, and conclude that the exam will not be on Friday, otherwise it would not be unexpected, and inductively that it cannot be on any other day of the week either. As a result, they do not study for the exam and, surely enough, a totally unexpected exam is given on Monday."

If we let $\varepsilon$ denote "there will be an exam today," then the flaw in the students' reasoning is that, while it is possible for the sentence $\varepsilon \land \neg K_{i}\varepsilon$ to be true for a node (student) $n_i$, the same cannot possibly hold for the sentence $K_{i}(\varepsilon \land \neg K_{i}\varepsilon)$, so that there are limits to what is knowable to $n_i$.

In a distributed setting like the one we have been considering in this book, there is interest in generalizing the notion of individual knowledge embodied in the operator $K_i$ to notions of group knowledge, say by all the members of $N$. Two simple possibilities of generalization in this sense are summarized by the operators $S_N$ and $E_N$, intended respectively to convey the notions of knowledge by at least one node and by all nodes. In other words,

$$S_N \mathcal{P} = \bigvee_{n_i \in N} K_{i}\mathcal{P},$$

and

$$E_N \mathcal{P} = \bigwedge_{n_i \in N} K_{i}\mathcal{P}.$$  

Another similar possibility of generalization is that of the notion of implicit knowledge by the group $N$. The meaning of implicit knowledge of $\mathcal{P}$ by $N$,
denoted $I_n\mathcal{P}$, is that $\mathcal{P}$ can be concluded from the individual knowledge that
the members of $N$ have. For example, if $\mathcal{Q}$ and $\mathcal{Q} \rightarrow \mathcal{P}$ are both true
sentences, and moreover both $K_n\mathcal{Q}$ and $\mathcal{P} \rightarrow \mathcal{Q}$ hold for $n, n_j \in N$, then
$\mathcal{P}$ is also true and $I_n\mathcal{P}$ holds as well.
Associated with this notion of implicit knowledge is a notion of conservation, which
states that no communication can change the implicit knowledge of propositional
nature that $N$ has. This result, which we do not investigate in any further depth in
this book, is to be regarded with care. In particular, the requirement that the
conserved implicit knowledge be propositional is crucial, as otherwise the
conservation need not hold. For example, if for a true sentence $\mathcal{P}$ it holds that
$\neg K_n\mathcal{P}$ and $K_n\mathcal{P}$ for $n, n_j \in N$ such that $(n, n_j) \in E$, then a message sent by $n_j$
to $n$ containing $\mathcal{P}$ suffices for $K_n\mathcal{P}$ to become implicit knowledge, although
such was not the case prior to the receipt of the message by $n$. However, the
sentence $K_n\mathcal{P}$ is not propositional, and then no contradiction to the conservation
principle is implied by this acquisition of new implicit knowledge.
The next step in generalizing the notion of individual knowledge to broader notions
of group knowledge is to consider the notion of common knowledge by $N$. This
notion, denoted by $C_n\mathcal{P}$, for a true sentence $\mathcal{P}$, is such that
$C_n\mathcal{P} \equiv \bigwedge_{k \geq 1} E_N^{k}\mathcal{P}$,
where
$E_N^{1}\mathcal{P} = E_N\mathcal{P}$
and, for $k > 1$,
$E_N^{k}\mathcal{P} = E_N\mathcal{P} \wedge E_N^{k-1}\mathcal{P}$.
This notion is very hard to grasp
intuitively, but nonetheless it should be clear that
$C_n\mathcal{P} \rightarrow K_{i_1} \ldots K_{i_z}\mathcal{P}$
holds for any set of integers $\{i_1, \ldots, i_z\} \supset \{1, \ldots, n\}$ with $z \geq 1$. Another anecdote is
usually very helpful in building up some intuition on the notion of common
knowledge.
"A group of boys are playing together and have been advised by their parents
that they should not get dirty. However, it does happen that some of them, say
$k \geq 1$, get dirty, but only on their foreheads, so that no boy knows whether his
own forehead is dirty though he can see the others'. One of the parents then
shows up and states, 'At least one of you has a dirty forehead,' thereby
expressing a fact already known to all the boys if $k > 1$. The parent then asks
repeatedly, 'Can anyone prove that his own forehead is dirty?' If we assume
that all the boys are unusually intellectually gifted, and moreover that they all
reply simultaneously at each repetition of the parent's question, then every boy
replies 'No' to the first $k - 1$ questions, and the boys with dirty foreheads reply
'Yes' to the $k$th question."
What supports the boys' reasoning in replying to the parent's repeated questions is
the following inductive argument. If $k = 1$, then the only boy with a dirty forehead
replies "Yes" immediately upon the first question, because he knows that at least one boy has a dirty forehead, and seeing no one else in that condition he must be the one. If we inductively hypothesize that the boys reason correctly for \(1 \leq k \leq k'\) with \(k' \geq 1\), then for \(k = k' + 1\) we have the following. A boy with a dirty forehead sees \(k'\) other boys with dirty foreheads, while a boy with a clean forehead sees \(k' + 1\) boys with dirty foreheads. By the induction hypothesis, a boy with a dirty forehead must reply "Yes" to the \(k\)th question, because if he did not have a dirty forehead the other \(k'\) boys with dirty foreheads that he sees would all have replied "Yes" upon hearing the previous question. Because they did not, his own forehead must be dirty.

In the context of this anecdote, the issue of knowledge comes in as follows. If \(P\) represents the parent's statement concerning the existence of boys with dirty foreheads and \(N\) is the set of boys, then, before the statement, \(E_{N}^{k-1}P\) holds but \(E_{N}^{k}P\) does not. What the parent's statement does is to establish \(C_{N}P\), therefore \(E_{N}^{k}P\), which is the necessary state of knowledge for the boys' reasoning to be carried out.

The various notions of knowledge we have encountered so far relate to each other hierarchically in such a way that

\[
C_{N}P \rightarrow \cdots \rightarrow E_{N}^{p}P \rightarrow \cdots \rightarrow E_{N}^{p}P \rightarrow E_{N}^{p-1}P \rightarrow E_{N}^{p-1}P \rightarrow \cdots
\]

holds for every \(k \geq 1\). While every information that is "built in" the nodes constitutes common knowledge, the acquisition of new common knowledge is far from trivial, unless some sort of "shared memory" can be assumed, as in the case of the anecdote we presented on the dirty-forehead boys (the parent's statement can be regarded as having been "written" into such a shared memory). To see why acquiring new common knowledge may be important, we consider yet another anecdote.

"Two divisions of an army are camped on the hills surrounding a valley, and in the valley is the enemy army. Both divisions would like to attack the enemy army simultaneously some time the next day, because each division individually is outnumbered by the enemies. Having agreed on no plan beforehand, the divisions' generals are forced to rely on forerunners to convey messages to each other. Forerunners must go through the enemy's camp with their messages, and then do it at night, although the risk of being caught still exists and in addition they may get lost. Given that normally one hour is enough for the trip, and that at this particular night the forerunners travel uneventfully through the enemy's camp and do not get lost, how long does it take for an agreement to be reached between the two generals?"

Clearly, what the two generals seek in this anecdote is common knowledge of an agreement. The reader must quickly realize, though, that such a state of knowledge cannot be attained. Indeed, unless communication is totally reliable (as we have implicitly been assuming) and the model of distributed computation is the synchronous model, no new common knowledge can ever be attained. However, the literature contains examples of how to attain new common knowledge in the asynchronous model with reliable communication by restricting the definition of common knowledge to special global states (cf. Section 3.1).

### 2.4 Exercises

1. **Show that**, if the ring is locally oriented, then **Algorithm A_Compute_f** can be modified so that every node receives exactly \(n\) arguments of \(f\) even if \(n\) is even.
2. Describe how to simplify Algorithm S_Locally_Orient if the determination of K is not required (that is, if the algorithm is to run for the maximum possible number of iterations).
3. Show how to modify Algorithm A_Compute_f so that it can be used to establish a local orientation on the ring (i.e., show how it should be changed to treat the B's instead of the b's).

1. Show that, if the ring is locally oriented, then Algorithm A_Compute_f can be modified so that every node receives exactly n arguments of f even if n is even. Describe how to simplify Algorithm S_Locally_Orient if the determination of K is not required (that is, if the algorithm is to run for the maximum possible number of iterations). Show how to modify Algorithm A_Compute_f so that it can be used to establish a local orientation on the ring (i.e., show how it should be changed to treat the B's instead of the b's).

2.5 Bibliographic notes
Readers in need of references on concepts from graph theory, for use not only in this chapter but throughout the book, may choose from a variety of sources, including some of the classic texts, like Harary (1969), Berge (1976), Bondy and Murty (1976), and Wilson (1979). The asynchronous and synchronous models introduced in Section 2.1 are pretty standard in the field, and can also be found in Lamport and Lynch (1990), for example. Algorithm A_FIFO, used as example in that section, is from Barbosa and Porto (1995).
The material on anonymous systems in Section 2.2 is based on Attiya and Snir (1985), which later appeared in revised form in Attiya, Snir, and Warmuth (1988). Further developments on the theme can be found in Attiya and Snir (1991), Bodlaender, Moran, and Warmuth (1994), Kranakis, Krizanc, and van den Berg (1994), and Lakshman and Wei (1994).
Readers seeking additional information on the notions related to knowledge can look for the survey by Halpern (1986), as well as the guide to the logics involved by Halpern and Moses (1992). The material in Section 2.3 is drawn from a variety of publications, which the reader may seek in order to deepen the treatment of a particular topic. The application of knowledge-related notions to problems in the context of distributed computations dates back to the first version of Halpern and Moses (1990) and to Lehmann (1984). In Halpern and Moses (1990), the reader will also find the definitions of implicit and common knowledge, as well as the argument for the impossibility of attaining common knowledge in the asynchronous model or under unreliable communication. Fischer and Immerman (1986) describe situations in which common knowledge can be attained in the asynchronous model if communication is totally reliable and in addition one is restricted to considering only some special global states. The anecdote involving students and the unexpected exam is from Lehmann (1984). The conservation of implicit knowledge is from Fagin and Vardi (1986).
Problems related to the agreement between generals of a same army can be found in Lamport, Shostak, and Pease (1982) and in Dwork and Moses (1990). Additional work on knowledge in distributed systems has appeared by Halpern and Fagin (1989), Fagin, Halpern, and Vardi (1992), Neiger and Toueg (1993), and van der Meyden (1994).

Chapter 3: Models of Computation

Overview
In this chapter, we return to the topic of computation models for distributed algorithms. We start where we stopped at the end of Section 2.1, which was devoted essentially to introducing the asynchronous and synchronous models of distributed computation. In that section, we also introduced, along with examples throughout Chapter 2, Algorithms A_Template and S_Template, given respectively as templates to write asynchronous and synchronous algorithms.
Our first aim in this chapter is to establish a more detailed model of the distributed computations that occur under the assumption of both the asynchronous and the synchronous model. We do this in Section 3.1, where we introduce an event-based formalism to describe distributed computations. Such a formalism will allow us to be much more precise than we have been so far when referring to global timing issues in the
asynchronous case, and will in addition provide us with the necessary terminology to define
the time-related complexity measures that we have so far avoided.
This discussion of complexity measures appears in Section 3.2 where the emphasis is on
time-related measures for asynchronous algorithms, although we also discuss such
measures for synchronous algorithms and return to the issue of message complexity
introduced in Section 2.1.
We continue in Section 3.3 by returning to the template algorithms of Section 2.1 to provide
details on how asynchronous algorithms can be executed under the assumptions of the
synchronous model. In addition, we also indicate, but only superficially in this chapter, how
synchronous algorithms can be transformed into equivalent asynchronous algorithms.
Section 3.4 is dedicated to a deeper exploration of the synchronous model, which, as we
have indicated previously, although unrealistic possesses some conceptual and practical
features of great interest. Some of these are our subject in Section 3.4 as an example of a
computation that is strictly more efficient in time-related terms in the synchronous model than
in the asynchronous model, and another in which the initial assumption of full synchronism in
the process of algorithm design eventually leads to greater overall efficiency with respect to
existing solutions to the same problem.
Sections 3.5 and 3.6 contain exercises and bibliographic notes, respectively.

3.1 Events, orders, and global states
So far in the book there have been several occasions in which we had to refer to
global characteristics of the algorithms we studied and found ourselves at a loss
concerning appropriate conceptual bases and terminology. This has been most
pronounced in the case of asynchronous algorithms, and then we have resorted to
expressions as "concurrent", "scenario", and "causally affect" to make up for the
appropriate terminology and yet convey some of the intuition of what was actually
meant. This happened, for example, during our discussion of task migration in
Sections 1.6.2 and 2.1, in our introduction of the synchronous model in Section 2.1,
and in the proof of Theorem 2.4. As we indicated in Sections 2.1 and 2.3, such
imprecisions can be corrected easily once the appropriate concept of a global state
has been established. Such a concept lies at the core of our discussion in this
section.
The case of synchronous algorithms is clearly much simpler as far as the concepts
underlying global temporal issues are concerned. In fact, in describing Algorithms
S_Compute_AND and S_Locally_Orient, respectively in Sections 2.2.2 and 2.2.3,
we managed without any difficulty to identify the number of pulses that had to
elapse for termination of the algorithm at hand. This number, as we will see in
Section 3.2.1, essentially gives the algorithm's time-related measure of complexity,
which in the asynchronous case we have not even approached.
Our discussion in this section revolves around the concept of an event, and is
intended especially to the description of computations taking place in the
asynchronous model (that is, executions of asynchronous algorithms). However, as
we mentioned in Section 2.1, the conditions under which the synchronous model is
defined can be regarded as a particularization of the conditions for the
asynchronous model, and then all of our discussion is also applicable in its
essence to the synchronous model as well. We shall return to this issue later to be
more specific on how the characteristics of the synchronous model can be seen to
be present in our event-based formalism.
The concept of an event in our formalism is that of a fundamental unit of a
distributed computation, which in turn is an execution of a distributed algorithm. A
distributed computation is then viewed simply as a set of events, which we denote
by Ξ. An event ξ is the 6-tuple

\[ ξ = (n, t, ξ, \phi, \theta, \delta) \]

where

- \( n \) is the node at which the event occurs;
- \( t \) is the time, as given by \( n \)'s local clock, at which the event occurs;
is the message, if any, that triggered the event upon its reception by \( n_i \);
\( \sigma \) is the state of \( n_i \) prior to the occurrence of the event;
\( \sigma' \) is the state of \( n_i \) after the occurrence of the event;
\( \Phi \) is the set of messages, if any, sent by \( n_i \) as a consequence of the occurrence of the event.

This definition of an event is based on the premise that the behavior of each node during the distributed computation can be described as that of a state machine, which seems to be general enough. The computation \( \Xi \) then causes every node to have its state evolve as the events occur. We let \( \Sigma \) denote the sequence of states \( n_i \) goes through as \( \Xi \) goes on. The first member of \( \Sigma \) is \( n_i \)'s initial state. The last member of \( \Sigma \) (which may not exist if \( \Xi \) is not finite) is \( n_i \)'s final state.

This definition of an event is also general enough to encompass both the assumed reactive character of our distributed computations (cf. Section 1.4) and to allow the description of internal events, i.e., events that happen without any immediate external cause (understood as a message reception or the spontaneous initiation by the nodes in \( N_0 \), which ultimately can also be regarded as originating externally).

In order to be able to describe internal events and events associated with the spontaneous initiation by the nodes in \( N_0 \), we have allowed the input message associated with an event to be absent sometimes. The atomic actions that we have associated with asynchronous algorithms (cf. Algorithm \( A_{\text{Template}} \)) can then be regarded as sequences of events, the first of which triggered by the reception of a message (or corresponding to the spontaneous initial activity of a node in \( N_0 \)), and the remaining ones being internal events.

For synchronous algorithms, these definitions are essentially valid as well, but a few special characteristics should be spelled out. Specifically, because in the synchronous case it helps to assume that local computation within an interval of the global clock takes zero time (cf. Section 2.1) and because nodes in the synchronous case are in reality driven by the global clock and not by the reception of messages, at each node exactly one event can be assumed to take place at each pulse, with \( t \) being a multiple of an interval's duration. Such an event does not have an input message associated with it, because by assumption every message is in the synchronous model delivered in strictly less time than the duration of an interval. In addition to these events, others corresponding solely to the reception of messages may also happen, but then with a different restriction on the value of \( t \), namely that \( t \) be something else than a multiple of an interval's duration. Finally, internal events are now meaningless, because every event either has an input message associated with it, or occurs in response to a pulse, having in either case an external cause. The overall picture in the synchronous case is then the following. At the beginning of the first interval (i.e., at the first pulse), an event happens at each of the nodes in \( N_0 \). Subsequently, at each new pulse and at each node an event happens corresponding to the computation by that node on the messages (if any) that it received during the preceding interval. Other events may happen between successive pulses, corresponding exclusively to the reception of messages for use at the succeeding pulse. The reader should notice that this description of a synchronous computation is in entire accordance with Algorithm \( S_{\text{Template}} \), that is, the events for which \( t \) is a multiple of an interval's duration correspond to the actions in that algorithm. The other events do not correspond to any of the algorithm's actions, being responsible for establishing the sets \( MSG(s) \) for \( n_i \in N \) and \( s > 0 \).

Events in \( \Xi \) are strongly interrelated, as messages that a node sends in connection with an event are received by that node's neighbors in connection with other events. While this relationship is already grasped by the definition of an event, it is useful to elaborate a little more on the issue. Let us then define a binary relation, denoted by \( \prec \), on the set of events \( \Xi \) as follows. If \( \xi_1 \) and \( \xi_2 \) are events, then \( \xi_1 \prec \xi_2 \) if and only if one of the following two conditions holds.

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Both $\xi_1$ and $\xi_2$ occur at the same node, respectively at (local) times $t_1$ and $t_2$ such that $t_1 < t_2$. In addition, no other event occurs at the same node at a time $t$ such that $t_1 < t < t_2$.

ii. Events $\xi_1$ and $\xi_2$ occur at neighbor nodes, and a message exists that is sent in connection with $\xi_1$ and received in connection with $\xi_2$.

It follows from conditions (i) and (ii) that $<$ is an acyclic relation. Condition (i) expresses our intuitive understanding of the causality that exists among events that happen at the same node, while condition (ii) gives the basic cause-effect relationship that exists between neighbor nodes.

One interesting way to view the relation $<$ defined by these two conditions is to consider the acyclic directed graph $H = (\Xi, <)$. The node set of $H$ is the set of events $\Xi$, and its set of edges is given by the pairs of events in $<$. The graph $H$ is a precedence graph, and can be pictorially represented by displaying the events associated with a same node along a horizontal line, in the order given by $<$. In this representation, horizontal edges correspond to pairs of events that fall into the category of condition (i), while all others are in the category of condition (ii).

Equivalently, horizontal edges can be viewed as representing the states of nodes (only initial and final states are not represented), and edges between the horizontal lines of neighbor nodes represent messages sent between those nodes. Viewing the computation $\Xi$ with the aid of this graph will greatly enhance our understanding of some important concepts to be discussed later in this section and in Section 3.2.

The transitive closure of $<$, denoted by $<^+$, is irreflexive and transitive, and therefore establishes a partial order on the set of events $\Xi$. Two events $\xi_1$ and $\xi_2$ unrelated by $<^+$, i.e., such that

$$(\xi_1, \xi_2) \in \Xi \times \Xi - <^+$$

are said to be concurrent events. This denomination, as one readily understands, is meant to convey the notion that two such events are in no way causally related to each other.

In addition to its use in defining this concept of concurrent events, the relation $<^+$ can also be used to define other concepts of equally great intuitive appeal, as for example those of an event’s past and future. For an event $\xi$, we let

$${\text{Past}}(\xi) = \{ \xi' | \xi' \in \Xi \text{ and } \xi' <^+ \xi \}$$

and

$${\text{Future}}(\xi) = \{ \xi' | \xi' \in \Xi \text{ and } \xi <^+ \xi' \}.$$
These two sets can be easily seen to induce "conical" regions emanating from \( \xi \) in the precedence graph \( H \) and contain, respectively, the set of events that causally influence \( \xi \) and the set of events that are causally influenced by \( \xi \) (Figure 3.1). We now focus on a closer examination of the issues raised in the beginning of this section with respect to an appropriate conceptual basis and a terminology for the treatment of global timing aspects in a distributed computation. The key notion that we need is that of a consistent global state, or simply global state, or yet snapshot. This notion is based on the formalism we have developed so far in this section, and, among other interesting features, allows several global properties of distributed systems to be referred to properly in the asynchronous model. We will in this section provide two definitions of a global state. While these two definitions are equivalent to each other (cf. Exercise 1), each one has its particular appeal, and is more suitable to a particular situation. Our two definitions are based on the weaker concept of a system state, which is simply a collection of \( n \) local states, one for each node, and one edge state for each edge. If \( G \) is a directed graph, then the number of edge states is \( m \), otherwise it is \( 2m \) (one edge state for each of the two directions of each of the \( m \) edges). The state of node \( n \) in a system state is drawn from \( \Sigma_i \), the sequence of states \( n_i \) goes through as the distributed computation progresses, and is denoted by \( \sigma_i \). Similarly, the state of an edge \( (n_i \rightarrow n_j) \) is simply a set of messages, representing the messages that are in transit from \( n_i \) to \( n_j \) in that system state, i.e., messages that have been sent by \( n_i \) on edge \( (n_i \rightarrow n_j) \) but not yet received by \( n_j \). We denote this set by \( \Phi_{ij} \). The notion of a system state is very weak, in that it allows absurd global situations to be represented. For example, there is nothing in the definition of a system state that precludes the description of a situation in which a message has been sent by \( n_i \) on edge \( (n_i \rightarrow n_j) \), but nevertheless neither has arrived at \( n_j \) nor is in transit on \( (n_i \rightarrow n_j) \).

Our first definition of a global state is based on the partial order \( \prec^+ \) that exists on the set \( \Xi \) of events of the distributed computation, and requires the extension of \( \prec^+ \) to yield a total order, i.e., a partial order that includes exactly one of \( (\xi_1, \xi_2) \) or \( (\xi_2, \xi_1) \) for all \( \xi_1, \xi_2 \in \Xi \). This total order does not contradict \( \prec^+ \), in the sense that it contains all pairs of events already in \( \prec^+ \). It is then obtained from \( \prec^+ \) by the inclusion of pairs of concurrent events, that is, events that do not relate to each other according to \( \prec^+ \), in such a way that the resulting relation is indeed a partial order. A total order thus obtained is said to be consistent with \( \prec^+ \).

Given any total order \( < \) on \( \Xi \), exactly \( |\Xi| - 1 \) pairs \( (\xi_1, \xi_2) \in < \) can be identified such that every event \( \xi \neq \xi_1, \xi_2 \) is either such that \( \xi < \xi_1 \) or such that \( \xi_2 < \xi \). Events \( \xi_1 \) and \( \xi_2 \) are in this case said to be consecutive in \( < \). It is simple to see that, associated with every pair \( (\xi_1, \xi_2) \) of consecutive events in \( < \), there is a system state, denoted by \( \text{system\_state}(\xi_1, \xi_2) \), with the following characteristics.

- For each node \( n \), \( \sigma_i \) is the state resulting from the occurrence of the most recent event (i.e., with the greatest time of occurrence) at \( n_i \), say \( \xi_i \), such that \( \xi_i \prec \xi \) (this includes the possibility that \( \xi = \xi_i \)).
- For each edge \( (n_i \rightarrow n_j) \), \( \Phi_{ij} \) is the set of messages sent in connection with an event \( \xi \) such that \( \xi_i \prec \xi \) (including the possibility that \( \xi = \xi_i \)) and received in connection with an event \( \xi' \) such that \( \xi' \prec \xi_j \) (including the possibility that \( \xi' = \xi_j \)).
Figure 3.2: Part (a) of this figure shows a precedence graph, represented by solid lines, for $n = 2$. As $\prec$ is already transitive, we have $\prec^+ = \prec$. Members of $\prec^+$ are then represented by solid lines, while the dashed lines are used to represent the pairs of concurrent events, which, when added to $\prec^+$, yield a total order $\prec^+_t$ consistent with $\prec^+$. The same graph is redrawn in part (b) of the figure to emphasize the total order. In this case, system-state $(\xi_2, \xi)$ is such that $n_1$ is in the state at which it was left by the occurrence of $\xi_1$, $n_2$ is in the state at which it was left by the occurrence of $\xi_2$, and a message sent in connection with $\xi_2$ is in transit on the edge from $n_2$ to $n_1$ to be received in connection with $\xi_2$. Because $\prec^+_t$ is consistent with $\prec^+$, system_state $(\xi_2, \xi)$ is a global state, by our first definition of global states.

The first definition we consider for a global state is then the following. A system, state $\Psi$ is a global state if and only if either in $\Psi$ all nodes are in their initial states (and then all edges are empty), or in $\Psi$ all nodes are in their final states (and then all edges are empty as well), or there exists a total order $\prec^+_t$ consistent with $\prec^+$, in which a pair $(\xi_1, \xi_2)$ of consecutive events exists such that $\Psi = \text{system}_\text{state}(\xi_1, \xi_2)$ (Figure 3.2).

Our second definition of a global state is somewhat simpler, and requires that we consider a partition of the set of events $\Xi$ into two subsets $\Xi_1$ and $\Xi_2$. Associated with the pair $(\Xi_1, \Xi_2)$ is the system state, denoted by $\text{system}_\text{state}(\Xi_1, \Xi_2)$, in which $\sigma_i$ is the state in which $n_i$ was left by the most recent event of $\Xi_1$ occurring at $n_i$, and $\Phi_{ij}$ is the set of messages sent on $(n_i \rightarrow n_j)$ in connection with events in $\Xi_1$ and received in connection with events in $\Xi_2$.

The second definition is then the following. A system state $\Psi$ is a global state if and only if $\Psi = \text{system}_\text{state}(\Xi_1, \Xi_2)$ for some partition $(\Xi_1, \Xi_2)$ of $\Xi$ such that

$$\text{Past}(\xi) \subseteq \Xi_1$$

whenever $\xi \in \Xi_2$. (Equivalently, we might have required the existence of a partition $(\Xi_1, \Xi_2)$ such that

$$\text{Future}(\xi) \subseteq \Xi_2$$

whenever $\Xi \in \Xi_2$.) For simplicity, often we refer to such a partition as the global state itself. Note that there is no need, in this definition, to mention explicitly the cases in which all nodes are either in their initial or final states, as we did in the case of the first definition. These two cases correspond, respectively, to $\Xi_1 = \emptyset$ and $\Xi_2 = \emptyset$.

As we mentioned earlier, these two definitions of a global state are equivalent to each other. The first definition, however, is more suitable to our discussion in Section 5.2.1, particularly within the context of proving Theorem 5.5. The second definition, on the other hand, provides us with a more intuitive understanding of what a global state is. Specifically, the partition $(\Xi_1, \Xi_2)$ involved in this definition can be used in connection with the precedence graph $H$ introduced earlier to yield the following interpretation. The partition $(\Xi_1, \Xi_2)$ induces in $H$ a cut (a set of edges) comprising edges that lead from events in $\Xi_1$ to events in $\Xi_2$ and edges from
events in $\Xi_2$ to events in $\Xi_1$. This cut contains no edges from $\Xi_2$ to $\Xi_1$ if and only if $\text{system\_state}(\Xi_1, \Xi_2)$ is a global state, and then comprises the edges that represent the local states of all nodes (except those in their initial or final states) in that global state, and the edges that represent messages in transit in that global state (Figure 3.3).

We also mentioned at the beginning of this section that our discussion would apply both under full asynchronism and under full synchronism. In fact, when defining an event we explicitly described how the definition specializes to the case of full synchronism. It should then be noted that the concept of a global state is indeed equally applicable in both the asynchronous and the synchronous models, although it is in the former case that its importance is more greatly felt. In the

![Figure 3.3: Parts(a) and (b) show the same precedence graph for n= 2. Each of the cuts shown establishes a different partition (\(\Xi_1, \Xi_2\)) of \(\Xi\). The cut in part (a) has no edge leading from an event in \(\Xi_2\) to an event in \(\Xi_1\), and then \text{system\_state}(\Xi_1, \Xi_2)\) is a global state, by our second definition. In this global state, \(n_1\) is in its initial state, \(n_2\) is in the state at which it was left by the occurrence of \(\xi_2\), and a message is in transit on the edge from \(n_2\) to \(n_1\), sent in connection with \(\xi_2\), and to be received in connection with \(\xi_3\). The cut in part (b), on the other hand, has an edge leading from \(\xi_2 \in \Xi_2\) to \(\xi_3 \in \Xi_1\), so \text{system\_state}(\Xi_1, \Xi_2)\) cannot be a global state.

synchronous case, many global states can be characterized in association with the value of the global clock, as for example in "the global state at the beginning of pulse $s \geq 0$." However, there is nothing in the definition of a global state that precludes the existence in the synchronous case of global states in which nodes' local states include values of the global clock that differ from node to node.

Having defined a global state, we may then extend the definitions of the past and the future of an event, given earlier in this section, to encompass similar notions with respect to global states. If $\Psi$ is a global state, then we define its \textit{past} and \textit{future} respectively as
and

\[
Past(\Psi) = \bigcup_{\xi \in \Xi_1} (\{\xi\} \cup Past(\xi)) = \Xi_1
\]

and

\[
Future(\Psi) = \bigcup_{\xi \in \Xi_2} (\{\xi\} \cup Future(\xi)) = \Xi_2,
\]

where \(\Psi = \text{system\_state}\ (\Xi_1,\Xi_2)\) (this definition demonstrates another situation in which our second definition of a global state is more convenient). Similarly, we say that a global state \(\Psi_1\) comes \textit{earlier} in the computation \(\Xi\) than another global state \(\Psi_2\) if and only if \(Past(\Psi_1) \subset Past(\Psi_2)\) (or, alternatively, if \(Future(\Psi_2) \subset Future(\Psi_1)\)).

This definition will be of central importance in our discussion of stable properties in Chapter 6. Another related definition, that of an earliest global state with certain characteristics in a computation, can be given likewise. We shall return to it in detail in Section 9.3.1.

In finalizing this section, the reader should return to our discussion at the section's beginning to recognize the importance of the concepts we have introduced in establishing a rigorous and meaningful terminology. In particular, it should be clear that the partial order \(\prec\) and the notion of a global state suffice to do away with all the ambiguities in our previous use of expressions like "concurrent," "scenario," and "causally affect" when referring to global timing aspects in the asynchronous model.

### 3.2 The complexity of distributed computations

Analyzing the complexity of any computation is a means of expressing quantitatively how demanding that computation is on the resources that it requires to be carried out. Depending on the type of computation one is considering, such resources may include the number of processor cycles, the number of processors, the number of messages that are sent, and various other quantities that relate to the resources upon which the computation's demands are heaviest. Determining which resources are crucial in this sense is then the fundamental issue when defining the appropriate measures of complexity for a computation. For example, for sequential computations the chiefest resource is time, as given by the number of processor cycles that elapse during the computation, but often the number of memory cells employed is also important.

Another example comes from considering parallel computations. Quite often the study of such computations is concerned with the feasibility of solving a certain problem on more than one processor so that the computation can be solved faster than on one single processor. In such cases, one of the fundamental resources continues to be the number of processor cycles, but now the number of processors is also important, because it is the interplay of these two quantities that establishes the overall efficiency of the resulting algorithm and also how that algorithm relates to its sequential counterpart. Models of parallel computation adopting measures of complexity related to these two types of resource include the PRAM (Parallel Random Access Machine), which is essentially a synchronous model of parallel computation on shared-memory cells, as well as other distributed-memory variants, also synchronous.

Whereas the models of parallel computation we just mentioned are geared towards the so-called data parallelism, the computations we treat in this book relate more closely to what is known as control parallelism, and then the approach to measure complexity needs to be substantially revised. Data parallelism is the parallelism of problem solving, that is, given a problem, the task is to solve it efficiently in parallel, which includes the design of an algorithm and the choice of a number of processors leading to the desired efficiency. Control parallelism, by contrast, is concerned with the computations that have to be carried out on a
fixed number of processors, interconnected in a fixed manner, like our graph \( G \). The computations of interest are not so much geared towards problem solving, but mainly towards controlling the sharing of resources, understood in a very broad sense, throughout the system. Very often this also includes the solution of problems very much in the data-parallel sense, but now the problem is stated on \( G \), which is fixed, so that the control-parallel aspects of the computation become far more relevant.

The complexity of distributed algorithms is based on the assumption that communication and time are the resources whose usage should be measured. Given this choice of crucial resources, the measures of complexity are expressed in the usual worst-case, asymptotic fashion, as functions of \( n \) and \( m \), respectively the number of nodes and edges in \( G \).

However, because \( G \) is in this book taken to represent a great variety of real-world systems (cf. Section 1.1) at some level of abstraction, some elaboration is required when establishing the appropriate complexity measures.

A convenient starting point to establish the complexity measures of distributed algorithms is to first consider communication as the predominant resource under demand. This does not mean that time ceases to be a relevant issue, but rather that only the time directly related to communication should be taken into account. This approach takes care of most of our needs in the book, and is our subject in Section 3.2.1. In Section 3.2.2, we relax this assumption that communication takes precedence over time that is not related to communication, and then the time that a node spends computing locally becomes a third resource whose usage is to be measured. The resulting extended definitions of complexity will be of use especially in Section 9.3.3.

### 3.2.1 Communication and time complexities

If communication is the dominating resource under demand, then the complexity of a distributed algorithm is expressed as two measures. The first measure is the already seen message complexity (cf. Section 2.1), which is given by the number of messages sent between neighbors during the computation in the worst case, that is, the maximum number of messages when variations in the structure of \( G \) are considered (when applicable), as well as all possible executions of the algorithm (each yielding a different set of events, in the terminology of Section 3.1). Alternatively, this measure can be substituted by the more accurate message-bit complexity, or simply bit complexity, which can be useful in conveying relevant differences among algorithms when the messages' lengths depend on \( n \) or \( m \) (as opposed to being \( O(1) \)). This is, for example, the case of Algorithm \( A_{\text{FIFO}} \) of Section 2.1 and of Algorithm \( S_{\text{Locally Orient}} \) of Section 2.2.3. In the former case, a message may contain a task's identification and a node's identification (it may also contain a migrating task, but we may assume for our present purposes that such a message does not actually contain the task's code, which would already be present at all nodes, but rather simply the task's identification), and then the algorithm's bit complexity is, by Theorem 2.3, \( O(m_k(\log|N| + \log n)) \). In the latter case, messages are sent containing integers no larger than \( n \) (cf. (2.16) and (2.17)), so that the algorithm's bit complexity is \( O(n \log^2 n) \). The bit complexity will be sometimes employed for our analyses in the book.

The other measure that contributes to expressing the algorithm's complexity is its time complexity, which, in very loose terms, is given by the time spent in communication that elapses during the computation in the worst case (again a maximum over the possible structures of \( G \) and over all of the algorithm's executions). Any further elaboration on this definition requires that we consider the asynchronous model and the synchronous model separately.

For the synchronous model, the assumption that the cost of communication dominates all others should come as no surprise, because since the definition of this model in Section 2.1, we have assumed that local computation takes no time (or takes a constant time, which can be assumed to be zero). Although we made this assumption so that the synchronous model could be described without much further elaboration, in this section the assumption comes in handy as well, because it is in full accord with the assumed dominance of communication costs.

The definition of the time complexity in the synchronous model is rather simple, and amounts essentially to counting the number of pulses that elapse during the computation. In essence, then, already in Sections 2.2.2 and 2.2.3, we would have been able to express the time
complexities of Algorithms \texttt{S\_Compute\_AND} and \texttt{S\_Locally\_Orient}, respectively. In fact, in Section 2.2.2 we saw that Algorithm \texttt{S\_Compute\_AND} requires $O(n)$ pulses for completion, and that is then its time complexity. Similarly, in Section 2.2.3 the number of iterations required by Algorithm \texttt{S\_Locally\_Orient} was seen to be given by $O(\log n)$, and because each iteration comprises $O(n)$ pulses, the time complexity of that algorithm is $O(n \log n)$.

Defining the time complexity for the synchronous model in this straightforward fashion may seem to the reader not to be in complete agreement with our stated purpose of measuring time solely as it relates to communication. After all, many pulses may elapse without any communication taking place (cf., for example, the synchronous algorithms presented in Sections 2.2.2 and 2.2.3), but such pulses do nevertheless get counted when assessing the algorithm's time complexity. What should be considered to resolve this apparent conflict is that, as we have mentioned more than once already, in the synchronous model messages are as important as their absence. By including in the time complexity intervals during which messages are not sent, we are essentially accounting for the time needed to convey information through the absence of communication as well.

In the asynchronous model, the assumption that the time complexity only takes into account the time to perform communication leads to the following methodology to compute an algorithm's time complexity. First assume, as in the synchronous model, that local computation takes no time, and also that the time to communicate one message to each node in a nonempty subset of a node's set of neighbors is $O(1)$.

The time complexity is then the number of messages in the longest causal chain of the form "receive a message and send a message as a consequence" occurring in all executions of the algorithm and over all applicable variations in the structure of $G$.

This definition can be made more formal, but before we do that let us consider two important related issues. First of all, it should be clear that the time complexity can never be larger than the message complexity, because every message taken into account to compute the former is also used in the computation of the latter. The usefulness of the time complexity in spite of this relationship with the message complexity is that it only considers messages that happen "sequentially" one after the other, that is, messages that are causally bound to one another. Essentially, then, the time complexity in the asynchronous case can be regarded as being obtained from the message complexity by trimming off all the messages that are "concurrent" to those in the longest receive-send causal chain.

The second issue is that the assumption of $O(1)$ message transmission times for the computation of the time complexity is only completely valid if every message has length $O(1)$ as well. However, we do maintain the assumption to compute the time complexity even otherwise, because taking variable lengths into account would not contribute qualitatively to establishing what the lengthiest causal chain is. In addition, the effect of variable length is already captured by the algorithm's bit complexity, introduced earlier in this section, which should be used when needed.

The way to define the time complexity of an asynchronous algorithm more formally is to resort to the precedence graph $H$ introduced in Section 3.1. This graph summarizes the essential causal dependencies among events in the computation, and allows the definition of the time complexity to be given rather cleanly as follows. Let every edge in $H$ be labeled either with a 1, if it corresponds to a message, or with a 0, otherwise. Clearly, this reflects our assumptions that messages take constant time to be sent between neighbors and that local computation takes no time. The time complexity for fixed $G$ and $H$ (i.e., for a fixed execution of the algorithm) is then the length of the longest directed path in $H$, with the labels of individual edges taken as their lengths. Taking the maximum over the applicable variations of $G$ and over all the executions of the algorithm (all $H$s) yields the desired measure.

The reader should now be in position to return to the asynchronous algorithms given previously in the book, and have their time complexities assessed to $O(1)$, in the case of Algorithm \texttt{A\_FIFO}, and to $O(n)$, in the case of Algorithm \texttt{A\_Compute\_f}.

### 3.2.2 Local and global measures

Assuming that communication dominates the complexity of a distributed computation, and that in turn local computation takes no time, is reasonable for many of the systems discussed in Section 1.1, especially computer networks and networks of workstations. However, $G$ is
intended to model a greater variety of message-passing systems, and for some of these, including the multiprocessors also discussed in Section 1.1, such an assumption may be a bit too strong.

In this section, we consider the impact of facing nonconstant local processing times, and expand our collection of complexity measures to encompass others that may reflect this extended view more appropriately. In the synchronous case, all that would be required would be to let the duration of an interval of the global clock be a function of $n$ and $m$. This function would then yield a third complexity measure for the synchronous model, and everything else would remain essentially as is. It is interesting to note, however, that under this broader assumption on the duration of an interval the overall picture of a synchronous computation would change a little. Specifically, it would be possible to send messages at any point inside an interval, not only at the interval's beginning. Furthermore, in terms of the event-based formalism of Section 3.1, internal events would exist in the synchronous model as well.

It is important to note, in the synchronous model, that at pulse $s > 0$ a node $n$ may need to examine the set of $MSG(s)$ of messages received during the previous interval, and the time to do this should continue to be assumed constant even when the time to do local processing is taken to be variable. What this implies, together with the assumption we have made so far that a node may send one message to all of its neighbors in parallel, is that none of the synchronous algorithms we have seen so far requires taking local processing times to be anything else than constant. Even Algorithm $S_{Locally\_Orient}$, in which by (2.16) and (2.18) it would seem that the examination of $O(n)$ sets of messages is required at a single pulse, can be easily written in more detail and then seen to require the examination of only one such set per pulse.

In the context of this book, however, it is in the asynchronous case that nonconstant times for local computation will be most important, although not until Chapter 9. In the asynchronous model, then, we shall let the local time complexity refer to the time to perform local computation upon receiving a message. This is then the complexity of an atomic action in Algorithm $A_{Template}$. Whenever we use this complexity measure in the book, and if confusion may arise, we refer to the algorithm's time complexity as its global time complexity. Algorithm $A_{Compute\_f}$ is the only algorithm we have seen so far for which variable local processing times may need to be considered. What leads to this is that the computation of

$$f(b_1^1, \ldots, b_i^n)$$

in (2.11) may require a time to be performed that is a function of $n$. Considering this algorithm carefully leads us to other situations in which it would be justifiable to assume nonconstant local processing times. As we mentioned earlier in Section 3.2, often a distributed algorithm is designed to solve a problem that is posed on $G$. Typical examples of such problems are the ones in consider in Sections 4.2, and 4.3, and in Chapter 7, in which we discuss graph algorithms. Clearly, in such cases a possibility would be to have all nodes transmit their local share of information on the structure of $G$ to a previously designated node (a leader—cf. Section 5.1), which would then solve the problem locally and then after that possibly spread the solution to the other nodes. This would be very much in the style of Algorithm $A_{Compute\_f}$, although the assumed anonymity in the case of that algorithm disallows the existence of a leader altogether, as we discuss in section 5.1. In fact, in the presence of anonymity there is no other choice but to program all nodes to perform the same computation, as we discussed in Section 2.2.

However, if the system is not anonymous, then the alternative of coalescing all the information regarding $G$ into a leader for solution of the problem is a real possibility, and for this possibility it is important to consider the local time complexity of solving the entire problem in one single node. Moreover, concentrating all the relevant information in the leader may have $O(nm)$ message complexity (considering that each message contains a constant number of node identifications) and $O(n)$ time complexity, which, after added to the complexity of electing a leader, should also be compared with the corresponding measures elicited by the fully distributed alternative, in which all nodes participate in the solution of the problem by computing on its share of the problem's input (the structure of $G$).

### 3.3 Full asynchronism and full synchronism

Having introduced the complexity measures of relevance for distributed algorithms, in this section we return to the question, first raised in Section 2.1, of the
equivalence between the asynchronous and synchronous models. What we do is first to indicate explicitly how Algorithm $S_{\text{Template}}$ can be used to express an asynchronous algorithm (originally written over the template given by Algorithm $A_{\text{Template}}$). Then, conversely, we show how to employ Algorithm $A_{\text{Template}}$ as a basis to transform a synchronous algorithm (originally written over the template Algorithm $S_{\text{Template}}$) into an asynchronous algorithm. The first part is simpler, because an asynchronous algorithm runs under all possible variations in the timing of the asynchronous model, in particular in the variation that corresponds to the synchronous model. The only concern we must have when translating Algorithm $A_{\text{Template}}$ into Algorithm $S_{\text{Template}}$ is that, in the former, atomic actions are in general triggered by the arrival of messages and executed when the corresponding Boolean conditions hold, while in the latter nodes are driven solely by the global clock and operate on the sets of messages received during the preceding intervals. Algorithm $S_{\text{Template}}$ makes no provisions to condition the execution of an action upon the validity of a Boolean expression, which must then be treated inside the action itself. What this amounts to in the translation of an asynchronous algorithm into a synchronous one is that, upon the occurrence of pulse $s=1$, only those messages in $MSG(1)$ received by $n \in N$ on edges for which the corresponding Boolean conditions in the asynchronous algorithm hold can lead to the execution of the corresponding actions. The others must be held for reconsideration upon the occurrence of pulse $s=2$. In general, then, at pulse $s > 0$ a node $n_i$ may compute on messages from any of $MSG(1), \ldots, MSG(s)$, at which occasion those messages are deleted from the set to which they belong so that the remaining ones may be considered in further pulses. This strategy is reflected in Algorithm $A$-$to$-$S_{\text{Template}}$, given next. The message $msg$ that in Algorithm $A_{\text{Template}}$ triggers $n_i$'s action is in Algorithm $A$-$to$-$S_{\text{Template}}$ viewed as a variable, initially equal to $\text{nil}$.

Algorithm $A$-$to$-$S_{\text{Template}}$:

Variables:

$\text{msg} \equiv \text{nil}$;

Other variables used by $n_i$, and their initial values, are listed here.

Listing 3.1

Input:

$s=0$, $MSG(0) = \emptyset$.

Action if $n_i \in N$:

- Do some computation;
- Send one message on each edge of a (possibly empty) subset of $Out$.

Listing 3.2

Input:

$s > 0$, $MSG(1), \ldots, MSG(s)$ such that $\text{origin}(msg) = c_k \in In_i$ with $1 \leq k \leq |In_i|$ for $msg \in \bigcup_{r=1}^{s} MSG(r)$.

Action:
while there exist \( msg, r \in \{1, \ldots, s\} \), and \( k \in \{1, \ldots, |In|\} \) such that \( msg \in MSG(r) \) with origin \( (msg) = c_k \) and \( B_k \) do

begin

Let \( r' \) be the smallest such \( r \) and \( k' \) any such \( k \);

\( msg := msg \);

Remove \( msg \) from \( MSG(r') \);

Do some computation;

Send one message on each edge of a (possibly empty) sub-set of \( Out \).

end.

In Algorithm A-to-S_Template, (3.1) is identical to (2.8) in Algorithm S_Template, while (3.2) reflects the need to evaluate the appropriate Boolean conditions before the action corresponding to (2.2) in Algorithm A_Template can be executed. What (3.2) does at pulse \( s > 0 \) is to select from \( MSG(1), \ldots, MSG(s) \) one of the earliest messages (in the synchronous sense) for which the corresponding Boolean condition is true, and then to allow the corresponding action of the asynchronous algorithm to be executed. When edges are FIFO, then (3.2) has to be worked on a little so that the choice of \( r' \) and \( k' \) guarantees that \( msg \) is, of the messages to have arrived on \( c_k \) but not yet received, the one to have arrived first. The reason why this might not happen is that \( B_k \) might be false for the first message and not for some other arriving on the same edge. If this happened for all edges, then \( n_i \) should simply halt, thereby indicating an error in the design of the algorithm, just as Algorithm A_Template would.

Clearly, this translation of an asynchronous algorithm to run in the synchronous model does not change the algorithm’s message complexity. In addition, the reader should check carefully that the same holds for the time complexity, that is, the number of pulses that elapse before termination of the resulting synchronous algorithm is exactly the number of messages in the lengthiest causal chain during an execution of the asynchronous algorithm. In addition, even if local processing cannot be assumed to be instantaneous, the translation does not increase the amount of local computation that needs to be carried out, even though it would seem that the need to check so many sets of messages in (3.2) could require further local processing. Readily, these sets can be organized as \( |In| \) queues of messages at \( n_i \), that is, one per incoming edge. Then the work that has to be done at (3.2) is the same that in Algorithm S_Template \( n_i \) has to do, and this is assumed to take constant time even if local processing cannot be so assumed.

The other direction of transformation, namely to transform a synchronous algorithm into an asynchronous one, is not as immediate, and in this section we only touch the issue superficially. We return to the subject in Section 5.3 for the complete details. Naturally, the problem in this case is that the resulting asynchronous algorithm must only allow the action of \( n_i \) at pulse \( s > 0 \) to be executed when the set \( MSG(s) \) is available, and this is not immediate in the synchronous model. It seems apparent, then, that in the resulting asynchronous algorithms there has to be more communication among the nodes than in the synchronous algorithm, so that these sets of messages can be ensured to contain all the pertinent messages when they are used. This further increase in communication may then lead to a greater time complexity for the asynchronous algorithm when compared to the synchronous algorithm.

A template for the translation of a synchronous algorithm into an asynchronous algorithm is given next as Algorithm S-to-A_Template. This algorithm employs an integer variable \( s_i \geq 0 \) for \( n_i \in N \). This variable, initially such that \( s_i = 0 \), is used to keep track of the pulses of the synchronous algorithm. A Boolean function \( DONE \),
(s) is used to indicate whether \( n_i \) is ready to proceed to the execution of the action that the synchronous algorithm would execute at pulse \( s_i + 1 \) for \( s_i \geq 0 \) (determining what this function should do is then essentially our subject in Section 5.3). Finally, the sets \( MSG(s) \) for \( s \geq 0 \) that Algorithm \( S_{\text{Template}} \) employs are also variables of Algorithm \( S\text{-to-A}_{\text{Template}} \), initially empty sets.

**Algorithm \( S\text{-to-A}_{\text{Template}} \):**

Variables:

\[ s_i = 0; \]

\[ MSG(s) = \emptyset \text{ for all } s \geq 0; \]

Other variables used by \( n_i \), and their initial values, are listed here.

### Listing 3.3

**Input:**

\[ msg_i = \text{nil} \]

**Action if \( n_i \in N_c \):**

- Do some computation:
  - Send one message on each edge of a (possibly empty) subset of \( Out \).

### Listing 3.4

**Input:**

\[ msg \text{ such that } \text{origin}_i(msg) = c_k \in \text{In}, \text{ with } 1 \leq k \leq |\text{In}|. \]

**Action:**

- if DONE\((s_i)\) then
  - begin
    - \( s_i := s_i + 1; \)
    - Do some computation;
    - Send one message on each edge of a (possibly empty) sub-
      set of \( Out \).
  - end
- else
  - Add \( msg \) to \( MSG(s_i + 1) \) if appropriate.

In Algorithm \( S\text{-to-A}_{\text{Template}}, \) (3.3) is identical to (2.1) in Algorithm \( A_{\text{Template}}, \) while (3.4) indicates how the function DONE\((s_i)\) is to be used to ensure that \( s_i \) can be incremented and that the action that Algorithm \( S_{\text{Template}} \) would perform at pulse \( s_i + 1 \) by (2.9) can be executed. When (3.4) is executed and DONE\((s_i)\) turns out to be false, then \( msg \), the message that triggered the action, is added to \( MSG; \) (\( s_i + 1 \) if appropriate) \( msg \) may be a message unrelated to the synchronous algorithm, that is, one of the messages constituting the additional communication traffic that the transformation requires).

As we remarked earlier, transforming a synchronous algorithm into an asynchronous one may lead to increases in both the message complexity and the
time complexity with respect to the synchronous algorithm. On the other hand, as will become apparent from the material in Section 5.3, the complexity that results from assuming nonconstant local processing times remains unchanged.

3.4 The role of synchronism in distributed computations

So far in the book we have stressed more than once that the synchronous model is, in at least one important sense, more "powerful" than the asynchronous model. The justification behind this informal notion has been that, in the synchronous model, the absence of messages conveys information to nodes, while in the asynchronous model nothing like this happens. In fact, in Sections 2.2.2 and 2.2.3 we have given two synchronous algorithms, respectively Algorithms S-Compute-AND and S-Locally-Orient, whose message complexities are strictly lower than that of Algorithm A-Compute-f, which is an asynchronous algorithm that may be used for the same purposes as those two synchronous algorithms. Of course, Algorithms S-Compute-AND and S-Locally-Orient do not have the same generality of Algorithm A-Compute-f, and then it might be argued that the improvement in message complexity is a consequence of their single-purpose nature, rather than the result of exploiting the characteristics of the synchronous model. However, it should be simple for the reader to verify that the same particularizations would not lead to any improvements in message complexity under the asynchronous model.

The central question that we address in this section is whether the synchronous model can also yield improvements in the time complexity of some asynchronous algorithms. As we remarked in Section 3.3, designing an algorithm for the synchronous model and then transforming it into an asynchronous algorithm may lead to an increase in both the message and time complexities with respect to the synchronous algorithm. But this does not imply that an asynchronous algorithm designed "from scratch" (i.e., not as the result of a transformation from a synchronous algorithm) would not have better complexities than the synchronous algorithm. In order to address this issue, we discuss a problem for which every asynchronous algorithm must have a strictly greater time complexity than a very straightforward synchronous algorithm that solves the same problem, thereby answering our question affirmatively.

The problem that we discuss is stated in very abstract terms, and is related to synchronization issues in distributed systems, although one will probably not easily find any practical situation to which it may be readily applicable. Stating the problem requires the introduction of the following new terminology. A port is a special edge in the graph \( G \), and a port event is an event that involves the sending of a message on a port. A node at which a port event may happen (i.e., it may send messages on a port) is called a port node. A session is informally defined in terms of our terminology of Section 3.1 as a set of events including at least one port event for every port and "delimited" by two global states. More formally, if \( \Xi \) is the set of events representing a distributed computation, then \( S \subseteq \Xi \) is a session if and only if \( S \) includes at least one port event for every port and in addition two global states \( (\Xi_1, \Xi_4) \) and \( (\Xi_2, \Xi_3) \) exist such that

For integers \( \mu \) and \( \sigma \) such that \( 1 \leq \mu \leq m \) and \( \sigma \geq 1 \), the problem that we consider is called the \( (\mu, \sigma) \)-session problem, and asks that a graph \( G \) with \( \mu \) ports and a distributed algorithm on \( G \) be found such that every execution of the algorithm can be partitioned into at least \( \sigma \) sessions. In addition, the set of events associated with every execution of the algorithm is required to be finite and every port node is required to obtain the information that the \( \sigma \) sessions have occurred.

Solving the \( (\mu, \sigma) \)-session problem is in general very simple. For the synchronous model, the problem is solved by choosing \( G \) to be any graph with no more than one port per node, and the synchronous algorithm to be such that every port node sends a message on its port at each of the pulses \( s = 0, \ldots, \sigma - 1 \). The time complexity of this synchronous algorithm is \( O(\sigma) \).

In the asynchronous model, we can also solve the problem with the same time complexity, as follows. Choose \( G \) as in the synchronous case, except that all port nodes can send messages to one another as well. The asynchronous algorithm is such that every port node performs \( \sigma \) rounds of sending a message on its port and then sending a message to every other port node containing information that it has finished its participation in the current session. The next round is only performed after similar messages have been received from
all other port nodes. The reader should verify that the time complexity of this algorithm is $O(\sigma)$ if $N_0$ contains all port nodes.

The difficulty arises when in $G$ we place a constant bound $b$ on $|Out_i|$ (or $|Inc_i|$, in the undirected case), thereby limiting the number of messages that a node may send in a single action of Algorithm S-Template or Algorithm A-Template ($G$ is in this case said to be $b$-bounded). Clearly, this bound does not affect our proposed synchronous solution to the $(\mu, \sigma)$-session problem, but the asynchronous solution is no longer feasible within $O(\sigma)$ time, because the broadcast to all port nodes at the end of each round can no longer be achieved within $O(1)$ time. We give in Theorem 3.1 a lower bound on the time required by any asynchronous solution to the problem.

**Theorem 3.1.**
For $b \geq 1$, every asynchronous solution to the $(\mu, \sigma)$-session problem in which $G$ is $b$-bounded must be such that the corresponding asynchronous algorithm has time complexity of at least $\sigma - 1 + \log_b + 1 \mu - 1$.

**Proof:** Let $G$ be $b$-bounded, and consider an asynchronous solution to the $(\mu, \sigma)$-session problem consisting of $G$ and of an asynchronous algorithm. Let $\Xi$ be the set of events corresponding to an execution of this algorithm, and label every event $\xi$ with an integer $(\xi)$ obtained inductively as follows. If $\xi$ happens at $n_i \in N$ when $n_i$ is in its initial state, then let $(\xi) = 0$. If not, then let $(\xi) = (\xi') + 1$, where $\xi'$ is the event having the greatest label among the events $\xi'$ in connection with which at least one message is sent and in addition $\xi' <^* \xi$. Informally, this labeling of the events in $\Xi$ corresponds to attaching to each event the number of messages on which it depends causally. Because $\Xi$ is finite, every label is finite as well. Let $t$ be the greatest label over $\Xi$. Clearly, the time complexity of the algorithm is at least $t$.

Now let

$$K = \left\lfloor \frac{t}{\log_b + 1} \mu \right\rfloor$$

and partition $\Xi$ into the $K$ subsets of events $\Xi_1, \ldots, \Xi_K$, where, for $k = 1, \ldots, K$, $\xi \in \Xi_k$ is such that $(k - 1) \leq (\xi) \leq k - 1 \mu - 1$. Clearly, then, all of

$$(\Xi_1 \cup \cdots \cup \Xi_\ell, \Xi_{\ell+1} \cup \cdots \cup \Xi_K)$$

for $1 \leq \ell < K$ are global states, because of the way the labels $(\xi)$ were assigned and of the fact that no two sets of $\Xi_1, \ldots, \Xi_K$ have any event with the same value for $(\xi)$. The next step is to partition every $\Xi_k$ into the sets $\Gamma_k$ and $\Theta_k$ such that all of

$$(\Gamma_k \cup \Theta_k \cup \cdots \cup \Xi_k)$$

for $1 \leq k < K$ and

$$(\Xi_1 \cup \cdots \cup \Xi_{K-1} \cup \Gamma_K, \Theta_K)$$

are global states, and furthermore the following two conditions hold for a sequence of ports $e_0, \ldots, e_k$ (this sequence may contain the same port more than once).

i. $\Gamma_k$ does not contain any port event involving $e_{k-1}$.

ii. $\Theta_k$ does not contain any port event involving $e_k$.

This partitioning can be done for all $k = 1, \ldots, K$ inductively as follows. Pick $e_0$ to be any arbitrary port, and assume that $e_{k-1}$ has been defined. If a port exists that is not involved in any port event in $\Xi_k$, then let $e_k$ be that port, $\Gamma_k = \Theta_k$, and $\Theta_k = \Xi_k$, thereby satisfying conditions (i) and (ii). If, on the other hand, every port is involved in at least one port event in $\Xi_k$, then let $\Xi_k$ be the earliest port event involving $e_{k-1}$ in $\Xi_k$, and consider the number of port events contained in the set

$$F_k = (\{\xi_1\} \cup Future(\xi_1)) \cap \Xi_k.$$
larger than the sum of the elements in the geometric progression of rate \( b + 1 \) starting at 1 and ending at
\[
(b + 1)^{\lfloor \log_{b+1} \mu \rfloor - 1-(k-1)\lfloor \log_{b+1} \mu \rfloor},
\]
that is,
\[
\frac{(b+1)^{\lfloor \log_{b+1} \mu \rfloor - 1-(k-1)\lfloor \log_{b+1} \mu \rfloor}}{b+1-1} = \frac{(b+1)^{\lfloor \log_{b+1} \mu \rfloor} - 1}{b} \leq \frac{\mu - 1}{b} \leq \mu - 1.
\]

What this means is that at least one of the \( \mu \) ports is not involved in any of the port events in \( F_k \). Taking one of these ports to be \( e_k \), \( \Gamma_k = \Xi_k - F_k \), and \( \Theta_k = F_k \) clearly satisfies conditions (i) and (ii). It can be easily verified that, in both cases, the resulting \( \Gamma_k \) and \( \Theta_k \) induce global states, as required (cf. Exercise 5).

By conditions (i) and (ii), the sets \( \Gamma_1, \Theta_{k-1} \cup \Gamma_k \) for \( 1 < k \leq K \), and \( \Theta_K \) cannot contain a session, because a session must include at least one port event for every port. What this amounts to is that every session must have a nonempty intersection with both \( \Gamma_k \) and \( \Theta_k \) for some \( k \) such that \( 1 \leq k \leq K \), meaning that \( K \) is the maximum number of sessions in \( \Xi \).

Because \( \Xi \) contains at least \( \sigma \) sessions, and considering the definition of \( K \), we have
\[
\sigma \leq K \leq \frac{t + 1}{\lfloor \log_{b+1} \mu \rfloor} + 1,
\]
and then \( t \geq (\sigma - 1) \lfloor \log_{b+1} \mu \rfloor - 1 \).

**Theorem 3.1** and our discussion earlier in this section indicate that the synchronous model possesses characteristics that allow synchronous algorithms to perform better than asynchronous algorithms with respect to the algorithms’ message and time complexities. In practice, however, the main interest in the synchronous model stems from the possibility of eventually obtaining an asynchronous algorithm from an algorithm originally designed for the synchronous model. This is the subject of extensive discussion in Section 5.3, but in this section we wish to highlight the role that this approach has had historically in the development of distributed algorithms.

We consider the problem of establishing a breadth-first numbering on the nodes of \( G \) when \( G \) is a directed graph. This problem asks that every node \( n_i \) be assigned a nonnegative integer \( d_i \) equal to the shortest distance from a designated node \( n_1 \) to \( n_i \) in terms of numbers of edges. Initially, \( a = \infty \) for all \( n_i \in N \), thereby taking care of those nodes to which no directed path exists from \( n_1 \). This problem is closely related to the problem of determining the shortest distances between all pairs of nodes when \( G \) is undirected (we treat this problem in Section 4.3).

Obtaining a synchronous algorithm to solve this problem is a trivial matter. At pulse \( s = 0, n_i \) sets \( d_i \) to zero and sends a message on every edge in \( Out_i \). For \( s > 0 \), if a node \( n_i \) receives at least one message during interval \( s - 1 \) and at pulse \( s \) it still holds that \( d_i = \infty \), then it must be that the shortest directed path from \( n_1 \) to \( n_i \) contains \( s \) edges. What \( n_i \) does in this case is to set \( d_i \) to \( s \) and then send a message on each edge in \( Out_i \). Readily, this algorithm requires no more than \( n - 1 \) pulses for completion and employs no more than \( m \) messages. Its time and message complexities are then, respectively, \( O(n) \) and \( O(m) \).

Historically, this simple synchronous algorithm has accounted for the introduction of an asynchronous algorithm of time complexity \( O(n \log n / \log k) \) and message complexity \( O(kn^2) \) for arbitrary \( k \) such that \( 2 \leq k < n \), while the best asynchronous algorithm available at the time had time complexity \( O(n^{2.5}) \) and message complexity \( O(n^{2.5}) \) for arbitrary \( t \) such that \( 0 < t \leq 0.25 \). The reader should experiment with these complexities to verify that, given any \( i \) in the appropriate range, there exists a \( k \), also in the appropriate range, such that the algorithm obtained from the synchronous algorithm is strictly better in at least one of the two complexity measures (and no worse in neither).

**3.5 Exercises**

1. Show that the two definitions of a global state given in Section 3.1 are equivalent to each other.
2. Obtain the time complexity of Algorithm A_FIFO when it is viewed as being executed over GP.
3. Give the details of an algorithm to concentrate upon one single node all the information on the structure of $G$. The resulting algorithm should have the complexities mentioned in Section 3.2.2.

4. Consider the $(\mu, \sigma)$-session problem in the asynchronous case, and suppose that a node does its $\sigma$ port events all before the broadcast. What is wrong with this approach?

5. Consider an event $\xi$ and the sets $\Xi_1 = \{\xi \cup \text{Past}(\xi)\}$ and $\Xi_2 = \{\xi\} \cup \text{Future}(\xi)$.

Show that both partitions $(\Xi_1, \Xi - \Xi_1)$ and $(\Xi - \Xi_2, \Xi_2)$ are global states.

6. Consider the synchronous algorithm for breadth-first numbering described in Section 3.4. Express that algorithm in the format given by AlgorithmS-Template.

### 3.6 Bibliographic notes

The material in Section 3.1 is based on Lamport (1978) and on Chandy and Lamport (1985), having also benefited from the clearer exposition of the concept of a global state to be found in Bracha and Toueg (1984). Additional insights into the concepts discussed in that section can be found in Yang and Marsland (1993), and in the papers in Zhonghua and Marsland (1994).

Formalisms different from the one introduced in Section 3.1, often with accompanying proof systems, have been proposed by a number of authors. These include temporal logic (Pnueli, 1981; Manna and Pnueli, 1992) and I/O automata combined with various proof techniques (Lynch and Tuttle, 1987; Chou and Gafni, 1988; Welch, Lamport, and Lynch, 1988; Lynch, Merritt, Weihl, and Fekete, 1994). Additional sources on related material are Malka and Rajsbaum (1992) and Moran and Warmuth (1993).

Most of the complexity measures introduced in Section 3.2 are standard in the field, and can also be looked up in Lamport and Lynch (1990). The reader may also find it instructive to check different models and associated complexity measures in the field of parallel and distributed computation. Source publications include Gibbons and Rytter (1988), Akl (1989), Karp and Ramachandran (1990), Feldman and Shapiro (1992), JáJá (1992), and Leighton (1992).

Section 3.3 is related to the discussion in Awerbuch (1985a), while the material in Section 3.4 is based mostly on the work by Arjomandi, Fischer, and Lynch (1983). The comments at the end of the section on the breadth-first numbering of nodes derive from Awerbuch (1985b).

Chapter 4: Basic Algorithms

**Overview**
Three basic problems are considered in this chapter, namely the problems of propagating information from a group of nodes to all nodes, of providing every node with information on which are the identifications of all the other nodes in $G$, and of computing the shortest distances (in terms of numbers of edges) between all pairs of nodes. Throughout this chapter, $G$ is an undirected graph.

The first problem is treated in Section 4.1, first in the context of propagating information from a group of nodes to all the nodes in $G$, and then in the context of propagating information from one single node to all others but with the additional requirement that the node originally possessing the information must upon completion of the algorithm have received news that all other nodes were reached by the propagation. Our discussion in Section 4.1 encompasses both the case of one single instance of the algorithm being executed on $G$ and of multiple concurrent instances initiated one after the other.

Section 4.2 contains material on the detection of $G$'s connectivity by all nodes in the form of providing each node with a list of all the other nodes in $G$. Although many algorithms can be devised with this end, the one we present builds elegantly on top of one of the algorithms discussed in the previous section, and is for this reason especially instructive.

Computing all-pair shortest distances is our subject in Section 4.3. This is the first graph problem we treat in detail in the book (others can be found in Chapter 7). Our approach in Section 4.3 is that of not only giving a fundamental distributed algorithm, but also providing a nontrivial example to be used to illustrate the relationship between the asynchronous and synchronous models of distributed computation when we further return to that topic in Section 5.3.

Sections 4.4 and 4.5 contain, respectively, exercises and bibliographic notes.

### 4.1 Information propagation

The problem that we consider in this section is that of propagating a piece of information (generically denoted by $inf$) from the single node (or group of nodes) that originally possesses it to all of $G$'s nodes. We divide our discussion into two parts. The first part is the presentation of algorithms to solve two important variations of the problem, and comes in Section 4.1.1.

The second part is a discussion of how to handle multiple concurrent instances of these algorithms without examining the contents of the message being propagated. This is the subject of Section 4.1.2.

#### 4.1.1 Basic algorithms

The problem of propagating information through the nodes of $G$ is the problem of broadcasting throughout $G$ information originally held by only a subset of $G$'s nodes. In this section, we consider two variations of this problem, known respectively as the Propagation of Information problem (or PI problem) and the Propagation of Information with Feedback problem (or PIF problem). In the PI problem, all that is asked is that all nodes in $G$ receive $inf$, whereas in the PIF problem the requirement is that not only all nodes receive $inf$ but also that the originating node be informed that all the other nodes possess $inf$. Let us discuss the PI problem first.

The PI problem can be solved by a wide variety of approaches, each one with its own advantages. For example, if $n_i$ is the node that originally possesses $inf$, then one possible approach is to proceed in two phases. In the first phase, a spanning tree is found in $G$. In the second phase, the spanning tree is employed to perform the broadcast, as follows. Node $n_i$ sends $inf$ on all the edges of the spanning tree that are incident to it. Every other node, upon receiving $inf$, passes it on by sending it on every edge of the spanning tree that is incident to it and is not the edge on which $inf$ arrived. Readily, if the spanning tree can be assumed to be available to begin with, then an asynchronous algorithm based on this strategy has message and time complexities both equal to $O(n)$.

This simple approach can be extended to the case in which $inf$ is originally possessed by more than one node. In this case, instead of a spanning tree we need a spanning forest, each of whose trees including exactly one of the nodes having $inf$ originally. The broadcast procedure on each tree is then entirely the same as we described for the case of a single initiator.
There are essentially three reasons why this simple and effective approach may be undesirable. The first reason is that the spanning tree (or spanning forest) may not be available to begin with and then has to be determined at the cost of additional message and time complexities. Of course, once the tree (or forest) has been determined, then it may in principle be used indefinitely for many further broadcasts from the same nodes, so that apparently at least the additional cost may be somehow amortized over many executions of the algorithm and then become negligible for most purposes. The second reason for investigating other approaches in spite of this possible amortization of the initial cost comes from considering possible applications of the algorithm. By forcing the broadcast to be carried out on the same edges all the time, we are in essence ignoring the effect the possible variations with time on the delays for message delivery and also ignoring the fact that relying on a single tree (or forest) may be unreliable. Although our models of computation make no provisions for either circumstance to actually be an issue, in practice it may certainly be the case.

The third reason for us to consider a different approach (and really the most important one) is that this other approach, although extremely simple, illustrates interesting principles of distributed algorithm design, and in addition constitutes a sort of foundation for the rest of this chapter and for other sections to come as well (Section 5.2.1, for example).

The solution we describe next to the PI problem does the broadcast by "flooding" the network, and for this reason has a higher message complexity than the one we gave based on a spanning forest. The very simple idea behind this approach is that all nodes possessing inf initially send it to all of their neighbors at the beginning (they all start concurrently). Every other node, upon receiving inf for the first time, sends it on to all of its neighbors, including the one from which it was received. As a result, a node receives inf from all of its neighbors. In this strategy, inf is propagated from the nodes that initially possess it as a "wave," and then reaches all nodes as fast as possible and regardless of most edge failures (that is, those that do not disconnect G). This figurative view of how the algorithm proceeds globally is quite helpful to one’s understanding of various distributed algorithms, including many that we discuss in this book, most notably in Chapter 5.6, and 9. This strategy is reflected in Algorithm A.PI, given next. In this algorithm, The set $N_0$ comprises the nodes that possess inf initially. A node $n_i$ employs the the Boolean variable reached (equal to false, initially) to indicate whether $n$ has been reached by inf. Node $n_i$, upon receiving inf from a neighbor, must check this variable before deciding whether inf should be passed on or not.

**Algorithm A.PI**

Variables:
- reached $= \text{false}$.

**Listing 4.1**

Input:
- $msg_i = \text{nil}$.

Action if $n_i \in N_0$
- reached $:= \text{true}$;
- Send inf to all $n_j \in \text{Neig}$.

**Listing 4.2**

Input:
count = inf.

**Action:**

  if not reached, then
  begin
  reached := true;
  Send inf to all \( n_j \in \text{Neig}_i \)
  end.

It should be instructive for the reader to briefly return to the interpretation of the functioning of this algorithm as a wave propagation to verify the following. It is impossible for a node \( n_k \) for which \( \text{reached} = \text{true} \) to tell whether a copy of inf it receives in (4.2) is a response to a message it sent in (4.1) or (4.2), or a copy that was already in transit to it when a message it sent out arrived at its destination. In the latter case, letting \( n_i \) be the node from which this copy of inf originated, the edge \((n_i, n_j)\) is one of the edges on which two waves meet, one from each of two members of \( N_0 \) (possibly \( n_i, n_j \), or both, depending on whether they belong to \( N_0 \)).

Because in \( G \) there exists at least one path between every node and all the nodes in \( N_0 \), it is a trivial matter to see that inf does indeed get broadcast to all nodes by Algorithm A_PI. In addition, by (4.1) and (4.2), it is evident that the message complexity of this algorithm is \( O(m) \) (exactly one message traverses each edge in each direction, totaling \( 2m \) messages) and that its time complexity is \( O(n) \).

Let us now consider the PIF problem. Unlike the PI problem, the PIF problem is stated only for the case in which inf is initially possessed by a single node. Similarly to the PI problem, a solution based on a spanning tree can also be adopted, having essentially the same advantages and drawbacks as in the case of that problem. In such a solution, \( n_1 \), the only node originally possessing inf, is viewed as the tree’s root, while every other node possesses a special neighbor, called parent, at node \( n_i \) on the tree path from \( n_i \) to \( n_1 \). The algorithm initiates with \( n_1 \) sending inf on all tree edges incident to it. Every other node \( n_i \), upon receiving inf from a neighbor for the first time, sets parent to be that neighbor and, if not a leaf, forwards inf on all tree edges incident to it, except the one leading to parent. If \( n_i \) is a leaf, then it sends inf back to parent, immediately upon receiving it for the first time. Every other node, except \( n_1 \), having received inf on every tree edge, sends inf to parent. Upon receiving inf on all tree edges incident to it, \( n_i \) has the information that inf has reached all nodes.

Clearly, this solution has both the message and time complexities equal to \( O(n) \)

The solution by flooding to the PIF problem that we now describe in detail is an extension of the flooding solution we gave in Algorithm A_PI to the PI problem. Similarly to the spanning-tree-based solution we just described, a variable parent is employed at each node \( n_i \) to indicate one of \( n_i \)'s neighbors. In contrast with that solution, however, this variable is no longer dependent upon a preestablished spanning tree, but rather is determined dynamically to be any of \( n_i \)'s neighbors as follows. When \( n_i \) receives inf for the first time, parent is set to point to the neighbor of \( n_i \) from which it was received. The algorithm is started by \( n_1 \), which sends inf to all of its neighbors. Every other node \( n_i \), upon receiving inf for the first time, sets parent appropriately and forwards inf to all of its neighbors, except parent. Upon receiving a copy of inf from each of its neighbors, \( n_i \) may then send inf to parent, as well. Node \( n_i \) obtains the information that all nodes possess inf upon receiving inf from all of its neighbors.

This algorithm is given next as Algorithm A-PIF. The variable parent is initialized to nil for all \( n_i \in N \). Node \( n_i \) also employs the variable count, initially equal to zero, to register the number of copies of inf received, and the Boolean variable reached, initially set to false, to indicate whether \( n_i \) has been reached by inf. Note that count = 0 if reached = false, but not conversely, because reached must become true right at the algorithm's onset, at which time count = 0. The set \( N_0 \) now comprises one single element, namely the node that initially possesses inf, so \( N_0 = \{n_1\} \).
Algorithm A_PIF:

Variables:
\[\text{parent}_i = \text{nil};\]
\[\text{count}_i = 0;\]
\[\text{reached}_i = \text{false}.\]

Listing 4.3

Input:
\[\text{msg}_i = \text{nil}.\]

Action if \(n_i \in N_0:\)
\[\text{reached}_i := \text{true};\]
Send inf to all \(n_i \in \text{Neig}_i.\)

Listing 4.4

Input:
\[\text{msg}_i = \text{inf} \text{ such that origin}(\text{msg}_i) = (n_i, n_i).\]

Action
\[\text{count}_i := \text{count}_i + 1;\]
if not reached, then
begin
\[\text{reached}_i := \text{true};\]
\[\text{parent}_i := n_i;\]
Send inf to every \(n_k \in \text{Neig}_i\) such that \(n_k \neq \text{parent}_i\);
end;
if \(\text{count}_i = |\text{Neig}_i|\) then
if \(\text{parent}_i \neq \text{nil}\) then
Send inf to \(\text{parent}_i.\)

It follows easily from (4.3) and (4.4) that the collection of variables \(\text{parent}_i\) for all \(n_i \in N\) establishes on \(G\) a spanning tree rooted at \(n_1\) (Figure 4.1). The leaves in this tree are nodes from which no other node receives \text{inf} for the first time. The construction of this tree can be viewed, just as in the case of Algorithm A_PIF, as a wave of information that propagates outward from \(n_1\) to the farther reaches of \(G\). Clearly, this construction involves \(|\text{Neig}_i| + \sum_{n_j \in \text{Neig}_i} (|\text{Neig}_i| - 1) = 2m - n + 1\) messages and \(O(n)\) time. If it can be shown that every edge on the tree (there are \(n-1\) such edges) carries an additional copy of \text{inf} from node \(n_i \neq n_1\) to \(\text{parent}_i\) by time \(O(n)\) as well, then the total number of messages involved in Algorithm

Figure 4.1: During an execution of Algorithm A_PIF, the variables \(\text{parent}_i\) for all nodes \(n_i\) are set so that a spanning tree is created on \(G\). This spanning tree is rooted at \(n_1\), and its leaves
correspond to nodes from which no other node received $inf$ for the first time. In this figure, a directed edge is drawn from $n_i$ to $n_j$ to indicate that $parent_i = n_j$.

$A_{PIF}$ is $2m = O(m)$, while its time complexity is $O(n)$. **Theorem 4.1** provides the necessary basis for this argument, with $T_i \subseteq N$ containing the nodes in the subtree rooted at node $n_i$.

**Theorem 4.1**

In Algorithm $A_{PIF}$, node $n_i \neq n$ sends $inf$ to parent, within at most $2d$ time of having received $inf$ for the first time, where $d$ is the number of edges in the longest tree path between $n_i$ and a leaf in $T_i$. In addition, at the time this message is sent every node in $T_i$ has received $inf$.

**Proof:** The proof proceeds by induction on the subtrees of $T_i$. The basis is given by $T_i$'s leaves, and then the assertion clearly holds, because no $n_i \in N$ is such that $parent_i$ is a leaf in $T$. Assuming the assertion for all the subtrees of $T_i$ rooted at nodes $n_j$, if $parent_i = n_j$, then $n_j$ sends $inf$ to $n_i$ within at most $2(d - 1)$ time of having received $inf$ for the first time. The theorem then follows by (4.3) and (4.4).

In addition to helping establish the complexity of Algorithm $A_{PIF}$, **Theorem 4.1** is also useful in polishing our view of the algorithm's functioning as a wave propagation. What happens then is that a wave is propagated forward from $n_i$, and then another wave is propagated ("echoed") back to $n_i$. This second wave is initiated concurrently at all the leaves of the spanning tree and collapses back towards $n_i$. Notice that the two waves are not really completely separated from each other. In fact, it may happen that the second wave reaches a node before the first wave has reached that node on all possible fronts (i.e., on all possible edges incident to that node).

**Corollary 4.2.**

In Algorithm $A_{PIF}$, node $n_i$ receives $inf$ from all of its neighbors within time $O(n)$ of having executed (4.3). In addition, at the time the last $inf$ is received every node in $N$ has received $inf$.

**Proof:** Immediate from **Theorem 4.1** applied to all nodes $n_i$ such that $parent_i = n_i$ and from (4.4).

Before ending this section, we wish to make one comment that relates the two algorithms we have studied to material we saw previously in Section 2.3. From the perspective of the material discussed in that section, Algorithms $A_{PI}$ and $A_{PIF}$ offer good examples of how the knowledge that the nodes have evolve as the algorithms are executed. In the case of Algorithm $A_{PI}$, before the algorithm is started it holds that $K_{n_i}^{P}$ for all $n_i \in N_i$, with

being any sentence that can be deduced from $inf$. When the algorithm is done, then $K_{n_i}^{P}$ holds for all $n_i \in N_i$.

The situation is quite similar for Algorithm $A_{PIF}$, although more can be said. Initially, it holds that $K_{n_i}^{P}$, and after the first wave has reached all nodes it holds that $K_{n_i}^{P}$ for all $n_i \in N_i$. In addition, by **Corollary 4.2**, when $n_i$ has received $inf$ from all of its neighbors it also holds that $K_{n_i}^{P}$ for all $n_i \in N_i$.

**4.1.2 Handling multiple concurrent instances**

Algorithms for propagating information throughout $G$ like the ones we discussed in the previous section are of fundamental importance in various distributed computations. Together with the three general techniques discussed in Chapter 5 (leader election, distributed snapshots, and network synchronization), these algorithms can be regarded as constituting fundamental building blocks for the design of distributed algorithms in general. In fact, algorithms for propagating information, either through all of $G$'s nodes (as in the previous section) or in a more restricted fashion, are themselves components used widely in the design of the other building blocks we just alluded to. Understandably, then, some of these algorithms have been incorporated in the design of communication processors as built-
in instructions to be executed by the nodes of $G$ when this graph represents a network of communication processors (cf. Section 1.2).

It is in this context that the question of how to handle multiple concurrent instances of Algorithms $A_{PI}$ and $A_{PIF}$ arises. In the case of Algorithm $A_{PI}$, multiple concurrent instances occur when the nodes in $N_0$ repeatedly broadcast a series of messages, say $inf_1, inf_2, \ldots$. A quick examination of the algorithm reveals that a possibility to handle such a series at a node $n_i$ is to employ a Boolean variable $\text{reached}_i^k$ in connection with $inf_k$, for $k \geq 1$. Upon arrival of a message, its contents indicate which variable to use. However, if $G$'s edges are FIFO, then another alternative can be considered that does not require an unbounded number of Boolean variables to be employed at each node, and furthermore does away with the need to inspect the contents of the messages (as befits a communication processor).

This alternative is based on the simple observation that, under the FIFO assumption, every node receives the stream of messages, on every edge incident to it, in the order the messages were sent by the nodes in $N_0$. The strategy is to employ $|\text{Neig}|$ counters at $n_i$ to indicate the number of messages already received on each of the edges in $\text{Inc}_i$. These counters, called $\text{count}_i^j$ for $n_j \in \text{Neig}$, are initially equal to zero and get incremented by 1 upon receipt of a message on the corresponding edge. In order to check whether such a message, when received from $n_j, n_j \in \text{Neig}$, is being received at $n_i$ for the first time, it suffices to check whether

$$\text{count}_i^j > \text{count}_i^l$$

for all $n_l \in \text{Neig}$ such that $j \neq l$. In the affirmative case, the message is indeed being received for the first time and should be passed on (cf. Exercise 2).

A similar question arises in the context of Algorithm $A_{PIF}$ when the stream of messages is sent by node $n_1$. As in the case of Algorithm $A_{PI}$, providing each node $n_i$ with an unbounded number of sets of variables, and then allowing $n_i$ to inspect the contents of incoming messages to decide which set to use, is an approach to solve the problem. Naturally, though, one wonders whether the FIFO assumption on the edges of $G$ can lead to a simplification similar to the one we obtained in the previous case. It should not be hard to realize, however, that the FIFO assumption does not necessarily in this case imply that the stream of messages is received at each node, on every edge incident to it, in the order it was sent by $n_1$, and then our previous strategy does not carry over (cf. Exercise 3). Nevertheless, the weaker assertion that every node is reached by the stream of messages in the order it was sent does clearly hold under the assumption of FIFO edges, but this does not seem to readily provide a solution that is independent of the messages’ contents.

### 4.2 Graph connectivity

The problem that we treat in this section is the problem of discovery, by each node in $N$, of the identifications of all the other nodes to which it is connected by a path in $G$. The relevance of this problem becomes evident when we consider the myriad of practical situations in which portions of $G$ may fail, possibly disconnecting the graph and thereby making unreachable from each other a pair of nodes that could previously communicate over a path of finite number of edges. The ability to discover the identifications of the nodes that still share a connected component of the system in an environment that is prone to such changes may be crucial in many cases. The algorithm that we present in this section is not really suited to the cases in which $G$ changes dynamically. The treatment of such cases requires techniques that are altogether absent from this book, where we take $G$ to be fixed and connected. The interested reader is referred to the literature for additional information. The algorithm that we present is not the most efficient one, either, but it is the one of our choice because it very elegantly employs techniques for the propagation of information seen in Section 4.1.1.

The algorithm is called Algorithm $A_{Test\_Connectivity}$, and its essence is the following. First of all, it may be started by any of the nodes in $N$, either spontaneously (if the node is in $N_0$) or upon receipt of the first message (otherwise). In either case, what a node $n_i$ does to initiate its participation in the
algorithm is to broadcast its identification, call it $id_i$, in the manner of Algorithm $A_{PIF}$. As we will see, this very simple procedure, coupled with the assumption that the edges in $G$ are FIFO, suffices to ensure that every node in $N$ obtains the identifications of all the other nodes in $G$.

The set of variables that node $n_i$ employs to participate in Algorithm $A_{Test Connectivity}$ is essentially an $n$-fold replication of the set of variables employed in Algorithm $A_{PIF}$, because basically what $n_i$ is doing is to participate in as many concurrent instances of Algorithm $A_{PIF}$ as there are nodes in $G$ (although not in the sense of Section 4.1.2, because now each instance is generated by a different node). So, for $n_i \in N$, 

- $\text{parent}_i^j$ (initialized to $\text{nil}$) indicates the node in $\text{Neig}_i$ from which the first $id_j$ has been received,
- $\text{count}_i^j$ (initially equal to zero) stores the number of times $id_j$ has been received, and the
- Boolean $(\text{equal to } \text{false},$ initially) is used to indicate whether $id_j$ has been received at least once. Another Boolean variable, $\text{initiated}$, initialized to $\text{false}$, is employed at $n_i$ to indicate whether $n_i \in N_0$. (Use of this variable is a redundancy, but we keep it for notational simplicity; in fact, $\text{initiated} = \text{true}$ if and only if there exists at least one $n_j \in N$ such that $\text{reached}_i^j = \text{true}$)

**Algorithm A_Test_Connectivity:**

**Variables:**

- $\text{parent}_i^k = \text{nil}$ for all $n_k \in N$;
- $\text{count}_i^k = 0$ for all $n_k \in N$;
- $\text{reached}_i^k = \text{false}$ for all $n_k \in N$;
- $\text{initiated}_i = \text{false}$.

**Listing 4.5**

**Input:**

$\text{msg}_i = \text{nil}$.

**Action if** $n_i \in N_0$:

- $\text{initiated}_i := \text{true}$;
- $\text{reached}_i^j := \text{true}$;

Send $id_i$ to all $n_j \in \text{Neig}_i$.

**Listing 4.6**

**Input:**

$\text{msg}_i = id_k$ such that $\text{origin}(\text{msg}_i) = (n_k, n_j)$ for some $n_k \in N$.

**Action:**

- if not $\text{initiated}_i$, then
begin
  \text{initiated}_i \assign \text{true};
  \text{reached}_i^i \assign \text{false};
end;

\text{count}_i^k := \text{count}_i^k + 1;

\text{if not} \begin{aligned}
\text{reached}_i^k & \text{ then} \\
\text{parent}_i^k & \assign n_i;
\end{aligned}
\begin{aligned}
\text{Send id}_i \text{ to all } n_j \in \text{Neig}_i; \\
\text{end};
\end{aligned}

\text{if} \begin{aligned}
\text{count}_i^k & \assign |\text{Neig}_i| \text{ then} \\
\text{parent}_i^k & \neq \text{nil} \text{ then}
\end{aligned}
\begin{aligned}
\text{Send id}_i \text{ to } \text{parent}_i^k;
\end{aligned}

In Algorithm \text{A\_Test\_Connectivity}, (4.5) and (4.6) should compared respectively with (4.3) and (4.4) of Algorithm \text{A\_PIF}. What this comparison reveals is that (4.3) and (4.5) are essentially the same, whereas (4.6) is obtained from (4.4) by the addition of the appropriate commands for \( n_i \) to initiate its participation in the computation if it is not in \( N_0 \).

As we mentioned earlier, this algorithm is based on the assumption that \( G \)'s edges are FIFO. To see that it works, it is helpful to resort to the pictorial interpretation as propagating waves that we employed in the previous section for the algorithms for information propagation. The wave that node \( n_i \) propagates forward with its identification reaches every other node \( n_j \) either when \( \text{initiated}_j = \text{true} \) or when \( \text{initiated}_j = \text{false} \). By (4.5) and (4.6), and because of the FIFO property of the edges, in either case \( id_i \) is only sent along the nodes on the path from \( n_j \) to \( n_i \) obtained by successively following the parent pointers after \( id_j \) has been sent on the same path. Therefore, by the time \( n_i \) receives \( id_j \) from all of its neighbors it has already received \( id_j \) at least once (cf. Exercise 4). Because this is valid for all \( n_j \in N_i \), then \( n_i \) must by this time know the identifications of all nodes in \( G \).

Algorithm \text{A\_Test\_Connectivity} can be regarded as the superposition of \( n \) instances of Algorithm \text{A\_PIF}, so its message complexity is \( n \) times the message complexity of that algorithm, that is, \( O(nm) \) (to be precise, each edge carries exactly \( n \) messages in each direction, so the total number of messages is \( 2nm \)). Because the lengths of messages depend upon \( n \), it is in this case appropriate to compute the algorithm's bit complexity as well. If we assume that every node's identification can be expressed in \( \lfloor \log n \rfloor \) bits, then the bit complexity of Algorithm \text{A\_Test\_Connectivity} is \( O(nm \log n) \). The time complexity of the algorithm is essentially that of Algorithm \text{A\_PIF}, plus the time for a node in \( N_0 \) to trigger the initiation of another node as far from it as \( n - 1 \) edges; in summary, \( O(n) \) as well.
4.3 Shortest distances

The last basic problem considered in this chapter is the problem of determining the shortest distances in $G$ between all pairs of nodes. Distances between two nodes are in this section taken to be measured in numbers of edges, so that the problem that we treat is closely related to the problem of breadth-first numbering that we considered briefly at the end of Section 3.4. The problem is now much more general, though, because in that section we concentrated solely on computing the distances from a distinguished node $n_1$ to the other nodes in $N$ that could be reached from it ($G$ was then a directed graph). In addition, in that section, $n_1$ was not required to know at the end of the algorithm the numbers that had been assigned to the other nodes.

Another requirement that we add to the algorithm to compute shortest distances is that at the end a node be informed not only of the distance from it to all other nodes, but also of which of its neighbors lies on the corresponding shortest path. Readily, the availability of this information at all nodes provides a means of routing messages from every node to every other node along shortest paths. When $G$ has one node for every processor of some distributed-memory system and its edges reflect the interprocessor connections in that system, this information allows shortest-path routing to be done (cf. Section 1.3).

We approach this problem by first giving a synchronous algorithm that solves it, and then indicating how the corresponding asynchronous algorithm can be obtained. The synchronous algorithm, called Algorithm $S_{\text{Compute Distances}}$, proceeds as follows. At pulse $s = 0$, every node sends its identification to all of its neighbors. At pulse $s = 1$, every node possesses the identifications of all nodes that are no farther from it than one edge (itself and its neighbors). A node then builds a set with the identifications of all those nodes that are exactly one edge away from it and sends this set to its neighbors. At pulse $s = 2$, every node has received the identifications of all nodes located no farther than two edges from it (itself, its neighbors, and its neighbors' neighbors). Because a node knows precisely which nodes are zero or one edge away from it, determining the set of those nodes that are two edges away is a simple matter. What happens then is that, in general, at pulse $s \geq 0$ a node sends to its neighbors a set containing the identifications of all those nodes that are exactly $s$ edges away from it. For $s = 0$, this set comprises the node’s own identification only. For $s > 0$, the set comprises every node identification received during interval $s - 1$, except those of nodes which are at most $s - 1$ edges away from itself. Clearly, no more than $n$ pulses are required. The last pulse may be an earlier one, though, specifically pulse $S$ if the set that the node generates at this pulse is empty. Clearly, all further sets the node generated would be empty as well, and then it may cease computing (although innocuous messages may still arrive from some of its neighbors). Naturally, the value of $S$ may differ from node to node. For simplicity, however, in the algorithm that we give next we let all nodes compute through pulse $s = n - 1$ (cf. Exercise 6).

As in Section 4.2, we let $id$ denote $n_i$’s identification. Variables used by Algorithm $S_{\text{Compute Distances}}$ are the following. The shortest distance from $n_i$, to $n_j \in N$ is denoted by $dist^j_i$, initially equal to $n$ (unless $j = i$, in which case the initial value is zero). The node in $Neig$ on the corresponding shortest path to $n_i \neq n$ is denoted by $first^j_i$, initially equal to $\text{nil}$. The set of identifications to be sent out to neighbors at each step is denoted by $set$; initially, it contains $n_i$’s identification only. In Algorithm $S_{\text{Compute Distances}}$, $N_0 = N$.

Algorithm $S_{\text{Compute Distances}}$:

Variables:

$dist^j_i := 0$

$dist^k_i := n$ for all $n_k \in N$ such that $k \leq i$;

$first^k_i := n\text{nil}$ for all $n_k \in N$ such that $k \leq i$;
Listing 4.7

```
set_i = {id_i}
```

Listing 4.8

```
Input:
0 < s ≤ n - 1, MSG(s) such that origin(set_i) = (n_i, n_j) for set_i ∈ MSG(s)
Action:
set_i := ∅
for all set_j ∈ MSG(s) do
  for all id_k ∈ set_j do
    if dist_i^k > s then
      begin
        dist_i^k := s;
        first_i^k := n_j;
        set_i := set_i ∪ {id_k}
      end;
    Send set_i to all n_k ∈ Neig.
```

Even before the correctness of Algorithm $S_{\text{Compute_Distances}}$ is established formally, evaluating its message and time complexities is a simple matter. If the algorithm functions correctly, then every node must receive the identification of every other node, and then by (4.7) and (4.8) every node’s identification must traverse every edge in both directions. If we take a node’s identification to be a message, then the number of messages employed by Algorithm $S_{\text{Compute_Distances}}$ is $2nm$, and its message complexity is then $O(nm)$. As in the case of Algorithm $A_{\text{Test_Connectivity}}$, message lengths are in this case dependent on $n$. If, as in the case of that algorithm, we assume that node identifications can be expressed in $\lceil \log n \rceil$ bits, then the bit complexity of Algorithm $S_{\text{Compute_Distances}}$ is $O(nm \log n)$. By the range of $s$ in (4.8), the time complexity of this algorithm is $O(n)$. What supports these results is Theorem 4.3.

**Theorem 4.3**

For $s \geq 0$ in Algorithm $S_{\text{Compute_Distances}}$, at pulse $s$ every node $n_i$ has received the identifications of exactly those nodes $n_j \in N$ such that the shortest paths between $n_i$ and $n_j$
contain no more than \( s \) edges. Furthermore, for \( j \neq i \), \( \text{dist}_i^j \) and \( \text{first}_i^j \) are, respectively, the number of edges and the neighbor of \( n_i \) on one such path.

**Proof:** The proof is by induction, and the basis, corresponding to pulse \( s = 0 \), is trivial. If we inductively assume the theorem’s assertion for pulse \( s - 1 \), then for pulse \( s > 0 \) we have the following. By the induction hypothesis, \( n_i \) has at pulse \( s - 1 \) received the identifications of all \( n_j \in N \) that are at most \( s - 1 \) edges away from it, and the corresponding \( \text{dist}_i^j \) and \( \text{first}_i^j \) have been set correctly. In addition, by the induction hypothesis and by (4.7) and (4.8), during interval \( s - 1 \) \( n_i \) has received from each of its neighbors the identifications of all \( n_j \in N \) that are \( s - 1 \) edges away from that neighbor. A node \( n_i \) is \( s \) edges away from \( n_i \) if and only if it is \( s - 1 \) edges away from at least one node in \( \text{Neig}_i \), so at pulse \( s \), \( n_i \) has received the identifications of all \( n_j \in N \) that are no more than \( s \) edges away from it. The theorem follows easily from the observation that, by (4.8), the variables \( \text{dist}_i^j \) and \( \text{first}_i^j \) for all \( n_j \in N \) that are \( s \) edges away from \( n_i \) are set when \( n_i \) first finds in \( \text{MSG}(s) \) the identification of \( n_j \).

Obtaining an asynchronous algorithm from Algorithm \( S\_\text{Compute\_Distances} \) goes along the lines of Section 3.3, where Algorithm \( S\text{-to-A}\_\text{Template} \) was given just for such purposes. We provide the result of such a transformation next, but only in Section 5.3.2, after we have discussed the general technique of synchronizers, will the reasons why the resulting asynchronous algorithm is correct be given. The asynchronous algorithm that we give to compute all the shortest distances in \( G \) is called Algorithm \( A\_\text{Compute\_Distances} \), and requires that all edges in \( G \) be FIFO edges (cf. Exercise 7). It is widely used, despite having been displaced by more efficient algorithms of great theoretical interest. In addition to its popularity, good reasons for us to present it in detail are its simplicity and the possibility that it offers of illustrating the synchronization techniques of Section 5.3.2.

In addition to the variables that in Algorithm \( S\_\text{Compute\_Distances} \) employs, in Algorithm \( A\_\text{Compute\_Distances} \) the following variables are also employed. For each \( n_j \in \text{Neig}_i \), a variable \( \text{level}_i^j \) is employed to indicate which sets of node identifications \( n_i \) has received from \( n_j \). Specifically, \( \text{level}_i^j = d \) for some \( d \) such that \( 0 \leq d < n \) if and only if \( n_i \) has received from \( n_j \) the identifications of those nodes which are \( d \) edges away from \( n_j \). Initially, \( \text{level}_i^j = -1 \). Similarly, a variable \( \text{state}_i \) is employed by \( n_i \) with the following meaning. Node \( n_i \) has received the identifications of all nodes that are \( d \) edges away from it for some \( d \) such that \( 0 \leq d < n \) if and only if \( \text{state}_i = d \). Initially, \( \text{state}_i = 0 \). Finally, a Boolean variable \( \text{initiated} \), initially set to \( \text{false} \), is used to indicate whether \( n_i \in N_0 \).

**Algorithm A\_Compute\_Distance:**

**Variables:**

\[
\begin{align*}
\text{dist}_i^i &= 0; \\
\text{dist}_i^k &= n \text{ for all } n_k \in N \text{ such that } k \neq i; \\
\text{first}_i^k &= \text{nil} \text{ for all } n_k \in N \text{ such that } k \neq i; \\
\text{set}_i &= \{id\};
\end{align*}
\]
\[ \text{level}_i^j = -1 \text{ for all } n_j \in \text{Neig}; \]
\[ \text{state}_i = 0; \]
\[ \text{initiated}_i = \text{false} \]

**Listing 4.9**

**Input:**
\[ \text{msg}_i = \text{nil} \]

**Action if** \( n_i \in N \): 
\[ \text{initiated}_i := \text{true}; \]
\[ \text{Send set}_i \text{ to all } n_j \in \text{Neig}. \]

**Listing 4.10**

**Input:**
\[ \text{msg}_i = \text{set}_j \text{ such that origin}(\text{msg}_i) = (n_i, n_j). \]

**Action:**
\[ \text{if not initiated}_i \text{ then} \]
\[ \text{begin} \]
\[ \text{initiated}_i := \text{true}; \]
\[ \text{Send set}_i \text{ to all } n_k \in \text{Neig}, \]
\[ \text{end}; \]
\[ \text{if state}_i < n - 1 \text{ then} \]
\[ \text{begin} \]
\[ \text{level}_i^j := \text{level}_i^j + 1; \]
\[ \text{for all } id_k \in \text{set}_i \text{ do} \]
\[ \text{if dist}_i^k > \text{level}_i^j + 1 \text{ then} \]
\[ \text{begin} \]
\[ \text{dist}_i^k := \text{dist}_i^k + 1; \]
\[ \text{first}_i^k := n_j; \]
\[ \text{end}; \]
\[ \text{if state}_i \leq \text{level}_i^j \text{ for all } n_j \in \text{Neig} \text{ then} \]
\[ \text{begin} \]
\[ \text{state}_i := \text{state}_i + 1; \]
\[ \text{set}_i := \{ id_k | n_k \in N \text{ and dist}_i^k = \text{state}\}; \]
\[ \text{Send set}_i \text{ to all } n_k \in \text{Neig}, \]
\[ \text{end} \]
\[ \text{end} \]
In Algorithm $A_{\text{Compute_Distances}}$, (4.9) and the portion of (4.10) that is executed only when $\text{initiated}=\text{false}$ are precisely the same as (4.7) in Algorithm $S_{\text{Compute_Distances}}$. The remainder of (4.10) corresponds to the translation of (4.8) into the asynchronous model. Although we relegate most of the discussion on the correctness of Algorithm $A_{\text{Compute_Distances}}$ to Section 5.3.2, in this section attention should be given to the fact that, if $\text{initiated}=\text{true}$, then (4.10) is only executed if $\text{state}_i < n - 1$. The point to notice is that this is in accord with the intended semantics of $\text{state}_i$, because if $\text{state}_i = n - 1$ then $n$ has already received the identifications of all nodes in $N$, and is then essentially done with its participation in the algorithm.

Another important point to be discussed right away with respect to Algorithm $A_{\text{Compute_Distances}}$ is that the FIFO property of edges, in this case, is essential for the semantics of the $\text{level}$ variables to be maintained. In (4.10), the distance from $n$ to $n_i$ is updated to $\text{level}_i^k + 1$ upon receipt of $\text{id}_k$, in a set from a neighbor $n_j$ of $n_i$ only because that set is taken to contain the identifications of nodes whose distance to $n_i$ is $\text{level}_i^k$. This cannot be taken for granted, though, unless $(n_i, n_j)$ is a FIFO edge.

The complexities of Algorithm $A_{\text{Compute_Distances}}$ can also be obtained right away. By (4.9) and (4.10), what node $n$ does is to send its identification to all of its neighbors, then the identifications of all of its neighbors get sent, then the identifications of all nodes that are two edges away from it, and so on. Thus $n$ sends $n$ messages to each of its neighbors, and the total number of messages employed is then $2nm$, yielding a message complexity of $O(nm)$ and a bit complexity of $O(nm \log n)$ if node identifications can be represented in $\lceil \log n \rceil$ bits.

The time complexity comes from considering that a node that is not in $N_0$ starts executing (4.10) within at most $n - 1$ time of the algorithm’s initiation, and that the longest causal dependency involving messages corresponds to sending a node’s identification as far as $n - 1$ edges away. The resulting time complexity is then $O(n)$.

Our treatment in Section 5.3 will provide a general methodology for assessing an asynchronous algorithm’s complexities from those of the synchronous algorithm from which it originated. As we mentioned in previous occasions, the natural expectation is that higher complexities arise in the asynchronous case, specifically to account for the additional number of messages and time consumed by the function $\text{DONE}_i$ appearing in (3.4). However, both the message and time complexities of Algorithm $A_{\text{Compute_Distances}}$ are exactly the same as its synchronous originator’s. The reason for this intuitively unexpected behavior will become clear in Section 5.3.2.

### 4.4 Exercises

1. Discuss what happens to Algorithm $A_{\text{PI}}$ if a node refrains from sending in $f$ to the neighbor from which it was received.
2. Write the algorithm that handles multiple concurrent instances of Algorithm $A_{\text{PI}}$ as suggested in Section 4.1.2.
3. Show, by means of an example, that FIFO edges do not suffice to guarantee that messages are received at all nodes in the order sent by node $n_1$, in the context of multiple concurrent instances of Algorithm $A_{\text{PI}}$.
4. Show, by means of an example, that FIFO edges do not suffice to guarantee, in Algorithm $A_{\text{Test_Connectivity}}$, that a node receives all the copies of every other node’s identification before receiving as many copies of its own identification as it expects.
5. Compare Algorithm $A_{\text{Test_Connectivity}}$ with the possibility of solving the problem by a leader (suppose such a leader already exists).
6. Modify Algorithm $S_{\text{Compute_Distances}}$ so that it terminates at a node when that node generates an empty list.
7. Show that Algorithm $A_{\text{Compute_Distances}}$ can do without the FIFO requirement and without the level variables, if lists are sent along with the distances to which they correspond.
1. Discuss what happens to Algorithm A_PI if a node refrains from sending \( inf \) to the neighbor from which it was received.
2. Write the algorithm that handles multiple concurrent instances of Algorithm A_PI as suggested in Section 4.1.2.
3. Show, by means of an example, that FIFO edges do not suffice to guarantee that messages are received at all nodes in the order sent by node \( n_i \), in the context of multiple concurrent instances of Algorithm A_PIF.
4. Show, by means of an example, that FIFO edges do not suffice to guarantee, in Algorithm A_Test_Connectivity, that a node receives all the copies of every other node's identification before receiving as many copies of its own identification as it expects. Compare Algorithm A_Test_Connectivity with the possibility of solving the problem by a leader (suppose such a leader already exists).
5. Modify Algorithm S_Compute_Distances so that it terminates at a node when that node generates an empty list.
6. Show that Algorithm A_Compute_Distances can do without the FIFO requirement and without the level variables, if lists are sent along with the distances to which they correspond.

4.5 Bibliographic notes


Chapter 5: Basic Techniques

Overview

This chapter expands considerably on the material of Chapter 4 by presenting three fundamental techniques that can be regarded as building blocks for distributed algorithms in general. These are the techniques of leader election, distributed snapshots, and network synchronization.

The problem of electing a leader in \( G \) is treated in Section 5.1, where we discuss various of the problem's characteristics and some of the successful approaches to solve it. Because this problem is intimately related with the problem of establishing a minimum spanning tree on \( G \), treated in Section 7.1, and Section 5.1 we introduce techniques that do not rely on spanning trees to elect a leader. In doing so, we first give an asynchronous algorithm for generic graphs, and then introduce two algorithms (one synchronous and one asynchronous) for the case in which \( G \) is a complete graph.

In Section 5.2, we introduce techniques to record, in a distributed fashion, a global state of an ongoing distributed computation. The ability to record global states is fundamental in several cases, and in the remainder of the book there will be several opportunities for us to employ this and related techniques, as for example in Sections 6.2 and 6.3. In Section 5.2, we give a general technique to record global states distributedly and also discuss some centralized variations of interest in some special contexts, as for example in our discussion of some methods of distributed simulation in Chapter 10.

Network synchronization is the subject of Section 5.3, in which we return to material previously covered in Section 3.3 to fill in the details of how to translate a synchronous algorithm into an asynchronous one. Our approach in this section is to provide the principles underlying the transformation, then to present a few techniques exhibiting different communication and time complexities, and then to discuss simplifications that apply in important special cases. Sections 5.4 and 5.5 contain exercises and bibliographic notes, respectively.

5.1 Leader election
A leader is a member of $N$ that all other nodes acknowledge as being distinguished to perform some special task. The leader election problem is the problem of choosing a leader from a set of candidates, given that initially a node $n_i$ is only aware of its own identification, denoted as previously by $id$. In the spirit of Section 2.2.1, it should be clear after some pondering that the leader election problem is meaningless in the context of anonymous systems. Moreover, even if the system is not anonymous, the leader election problem can only be solved for $G$ if every node’s identification is unique in $G$ (cf. Exercise 1), in which case the set of all identifications can be assumed to be totally ordered by $\prec$.

This assumption is fundamental in the approaches to leader election that take the leader to be the candidate with greatest identification. However, even if this is not the criterion, the ability to compare two candidates’ identifications is essential to break ties that may occur with the criterion at hand. (In fact, this is really why unique identifications are needed in the first place. In their absence, any criterion to select a leader from the set of candidates might deadlock for the absence of a tie breaker.) Another assumption that we make on a node’s identification is that it can be expressed in $\log n$ bits. Also, $G$ is throughout this section assumed to be an undirected graph.

The importance of electing a leader in a distributed environment stems essentially from the occurrence of situations in which some centralized coordination must take place in $G$, either because a technique to solve the particular problem at hand in a completely distributed fashion is not available, or because the centralized approach offers more attractive performance. Problems for which satisfactory techniques of a completely distributed nature are not available include the many recovery steps that have to be taken after $G$ undergoes a failure (or a topological change, in broader terms). A leader is in this case needed to coordinate, for example, the reestablishment of allocation and routing functions (if $G$ is organized to reflect a distributed-memory system). Although it has been our assumption throughout that $G$ is fixed, all the algorithms that we discuss in this book are also applicable to the cases in which $G$ varies if $G$ is guaranteed to remain constant for “sufficiently long.”

Examples to illustrate the importance of electing a leader when the centralized approach to a particular problem proves more efficient than the distributed one come from the area of graph algorithms, treated in Chapter 7 (although in that chapter we concentrate solely on problems for which efficient distributed solutions do exist). As we remarked in Section 3.2.2, such situations are in essence characterized by higher complexities for the distributed approaches than for coalescing into the leader information on the structure of $G$.

The leader election problem is very closely related to another problem that we treat in Section 7.1, namely the problem of establishing on $G$ a minimum spanning tree (or, as we discuss in that section, the problem of determining any spanning tree on $G$, which can be reduced to the former). Once a spanning tree has been established on $G$, a leader can be elected as follows. Every node assumes the role of $n_i$ in the PIF problem (cf. Section 4.1.1) and propagates a piece of information with feedback on the spanning tree. This piece of information is, if the node is a candidate, its identification. Otherwise, it is simply a token devoid of any special content. The way the $n$ propagations interact with each other is such that a node, before forwarding the information being propagated by any other node, must first ensure that its own information is propagated. If edges are FIFO, then by the time a node receives its own information from all of its neighbors on the tree it has also received the information that all other nodes propagated, and can then select as a leader the candidate with greatest identification. This procedure has message complexity of $O(n^2)$ (bit complexity of $O(n^2 \log n)$) and time complexity of $O(n)$.

Although more efficient approaches exist to elect a leader once a spanning tree has been established on $G$ (cf. Sections 7.1.1), the approach we just described is interesting because it hints immediately at a simple (although not very efficient, either) algorithm to elect a leader on a generic graph with FIFO edges. This algorithm is simply Algorithm A_TestConnectivity of Section 4.2 with $N_i$ being the set of candidates, slightly modified so that candidates broadcast their identifications, while the remaining nodes broadcast simply a token that serves the purpose of signaling to the node that already it has received every candidate’s identification. Upon receiving on all incident edges the information it propagated, a node (be it a candidate or otherwise) is then ready to choose as a leader the candidate with greatest identification. As in Section 4.2, this algorithm’s message complexity is $O(nm)$ ($O(nm \log n)$ bit complexity), and its time complexity is $O(n)$. 

A_Test_Connectivity
When $G$ is assumed to be some particular graph, these complexities must be revised accordingly. For example, if $G$ is a ring, then the algorithm's message complexity becomes $O(n^2)$, whereas if it is a complete graph the message complexity is $O(n^3)$. The resulting message complexity for a complete graph is particularly alarming, and it is to the problem of electing a leader on such a graph that we turn our attention now, aiming specifically at providing an algorithm of significantly lower message complexity.

We start with a synchronous algorithm, aiming at illustrating the technique more intuitively, and then provide an asynchronous algorithm. In both algorithms, $N_i$ is the set of candidates. The synchronous algorithm that we give is inspired in the following straight-forward synchronous algorithm to elect a leader on a complete graph. At pulse $s = 0$, every candidate sends its identification to all other nodes. At pulse $s = 1$, every node has received every candidate's identification and can then decide on a leader. The message complexity of this algorithm is $O(n^2)$, and its time complexity is $O(1)$. The synchronous algorithm that we derive from this has message complexity of $O(n \log n)$ and time complexity of $O(\log n)$, which have been proved optimal in the literature.

In order to decrease the message complexity from $O(n^2)$ to $O(n \log n)$, a candidate does not send its identification to all of its neighbors at the same pulse, but rather first communicates with one of its neighbors, then with two other neighbors, and so on. For $k \geq 1$, the $k$th set of neighbors with which a candidate communicates has size $2^{k-1}$, and then $\lfloor \log n \rfloor$ such sets have to exist to encompass all of the candidate's $n - 1$ neighbors (that is, $k \leq \lfloor \log n \rfloor$). When a candidate sends a neighbor a message, it is attempting to “capture” that neighbor, thereby becoming its "owner," so that the candidate that has captured all nodes at the end is the one to be chosen leader. A candidate succeeds in capturing a node if its identification is larger than those of the other candidates that are attempting to capture the same node at the same time, and larger than the identification of the node's current owner. A candidate only proceeds to attempting to capture the next subset of its neighbors if it succeeds in capturing all the neighbors it is currently attempting to capture. Otherwise, it ceases being a candidate.

The resulting synchronous algorithm, called Algorithm $S$ $\_ $ $E l e c t \_ L e a d e r \_ C$ ("C" for Complete), proceeds as follows. At an even pulse $s \geq 0$, a candidate $n_i$ sends a message $\text{capture}(id)$ to $2^{s/2}$ of those of its neighbors with which it still has not communicated. At an odd pulse $s > 0$, a node $n_i$ (candidate or otherwise) selects from those nodes that sent it a $\text{capture}$ message at pulse $s - 1$ the one with greatest identification. That node is then to become $n_i$'s owner if its identification is greater than $n_i$'s current owner's. If $n_i$'s owner changes, then it sends an $\text{ack}$ to its new owner. Only a candidate that receives as many $\text{ack}$'s as it sent $\text{capture}$'s remains being a candidate.

The following are the variables employed by $n_i$ in Algorithm $S$ $\_ $ $E l e c t \_ L e a d e r \_ C$. A Boolean variable $\text{candidate}_i$, initially set to false, indicates whether $n_i$ is a candidate. For each neighbor $n_j$ of $n_i$, a Boolean variable $\text{tried}_i^j$ (equal to false, initially) is used to indicate, if $n_i$ is a candidate, whether it has already attempted to capture $n_j$. Finally, $\text{owner}_id_i$ contains the identification of $n_i$'s owner. This variable's initial value is nil (we assume that nil < id for all $n_i \in N$).

**Algorithm $S$ $\_ $ $E l e c t \_ L e a d e r \_ C$:**

```
Variables:
\text{candidate}_i = false;
\text{tried}_i^j = false \text{ for all } n_j \in \text{Neig} ;
\text{owner}_id_i = \text{nil}.
```

*Definition (Gold [80]) (a) Let $M$ be a computable scientist and let $L$*
Input:
\[ s = 0, \text{MSG}(0) = \emptyset. \]

Action if \( n \in N_0 \):
- \( \text{candidate} := \text{true}; \)
- \( \text{owner} := \text{id}; \)
- Let \( n \) be a node in \( \text{Neig}_i \);
  \( \text{tried}_i := \text{true}; \)
- Send \( \text{capture(id)} \) to \( n_i \).  

---

### Listing 5.2

**Input:**
\[ s \text{ odd such that } 0 < s \leq 2^{\lceil \log n \rceil} - 1, \text{MSG}(s) \text{ such that } \ \text{origin}, (\text{capture(id)}) = (n_i, n_j) \text{ for } \text{capture(id)} \in \text{MSG}(s). \]

**Action:**
- Let \( n \in \text{Neig}_i \) be such that \( \text{id}_k \geq \text{id}_j \) for all \( \text{capture(id)} \in \text{MSG}(s); \)
  - if \( \text{owner}_id < \text{id}_k \) then
    - begin
      - if \( \text{candidate} \) then
        - \( \text{candidate} := \text{false}; \)
        - \( \text{owner}_id := \text{id}_k; \)
        - Send \( \text{ack} \) to \( n_k \)
    - end.

---

### Listing 5.3

**Input:**
\[ s \text{ even such that } 0 < s \leq 2^{\lceil \log n \rceil}, \text{MSG}(s). \]

**Action:**
- if \( \text{candidate} \) then
  - if \( |\text{MSG}(s)| < \min\{2^{(s/2)}, n - 2^{(s/2)}\} \) then
    - \( \text{candidate} := \text{false} \)
  - else
    - if \( s < 2^{\lceil \log n \rceil} \) then
      - begin
        - Let \( S \subset \text{Neig}_i \) be such that \( |S| = \min\{2^{s/2}, n - 2^{s/2}\} \) and \( \text{tried}_i \) = \text{false} for all \( n_j \in S; \)
        - \( \text{tried}_i := \text{true} \) for all \( n_j \in S; \)
        - Send \( \text{capture(id)} \) to all \( n_j \in S \)
      - end.
In Algorithm $S_{\text{Elect \_Leader \_C}}$, (5.1) and (5.3) correspond to opportunities that $n$ has to capture nodes if it is a candidate. The attempt to capture more nodes in (5.3) is conditioned upon having received as many ack’s as needed for the last attempt. In other words, if $s > 0$ is even, then the number of ack’s expected to be in MSG$(s)$ if $n_i$ is a candidate is $\min(n^{2^{k-1}}, n-2^{(s-2)/2})$ (this is the number of nodes $n_i$ attempted to capture at the previous even pulse). If MSG$(s)$ contains this number of ack’s, then $n_i$ sends capture’s to other $2^k$ nodes, unless the number of nodes which it still has not attempted to capture is less than this, in which case it must be $n_i = 2^{(s-2)/2} - 1$ (the expression in parentheses is the number of nodes it has captured so far). In (5.2), node $n$, decides whether to change its owner or not, regardless of whether it is a candidate. The node that $n_i$ considers to be its owner after all $2^{\lceil \log n \rceil} + 1$ pulses have elapsed (that is, the node whose identification is in $\text{owner\_id}$, at that time) is the elected leader from $n_i$’s standpoint. By (5.1) and (5.2), $n_i$’s owner is the node with greatest identification, and therefore from every node’s standpoint the elected leader is the same. Algorithm $S_{\text{Elect \_Leader \_C}}$ runs for $2^{\lceil \log n \rceil} + 1$ pulses. Because of the number of opportunities that a candidate tries to capture is at most $\lceil \log n \rceil$, the last pulse in (5.3) is only used for a candidate to process the last ack’s it has received, if any. The time complexity of Algorithm $S_{\text{Elect \_Leader \_C}}$ is, by (5.3), $O(\log n)$. The following theorem indicates how to assess the algorithm’s message complexity.

**Theorem 5.1.**
For $1 \leq k \leq \lceil \log n \rceil - 1$, the maximum number of nodes to reach pulse $s = 2k$ as candidates in Algorithm $S_{\text{Elect \_Leader \_C}}$ is $\lceil n/2^{k-1} \rceil$.

**Proof:** At pulse $s = 2k$, by (5.3) a node must have captured $2^{k-1}$ nodes to be still a candidate (i.e., it must have received $2^{k-1}$ ack’s). The assertion then follows from the fact that, by (5.2), any of the $n$ nodes may only be captured by at most one candidate at any even pulse. By Theorem 5.1, at pulse $s = 2^{\lceil \log n \rceil} - 2$ there may still be a number of candidates no greater than

$$\left\lceil \frac{n}{2^{\lceil \log n \rceil - 2}} \right\rceil \leq \frac{4n}{2^{\lceil \log n \rceil}} \leq \frac{4n}{n} = 4$$

so that the additional even pulse $s = 2^{\lceil \log n \rceil}$ is indeed needed for all but one of them to quit being a candidate.

**Corollary 5.2.**
Algorithm $S_{\text{Elect \_Leader \_C}}$ employs at most $2n \lceil \log n \rceil - n$ capture messages and at most $n \lfloor \log n \rfloor$ ack messages.

**Proof:** The initial number of candidates is at most $n$, so by (5.1) at pulse $s = 0$ at most $n$ capture’s are sent. For $1 \leq k \leq \lceil \log n \rceil - 1$, by (5.3) at pulse $s = 2k$ a candidate sends at most $2^k$ capture’s. By Theorem 5.1, the number of candidates at this pulse is no larger than $\lceil n/2^{k-1} \rceil$, and then the total number of capture’s is at most

$$\sum_{k=1}^{\lceil \log n \rceil} \frac{n}{2^k} = n \left( 1 + 2 + \ldots + 2^{\lceil \log n \rceil - 1} \right) = 3n \lceil \log n \rceil - n.$$ 

By (5.2), a node sends at most one ack per odd pulse, so that the total number of ack’s is no more than $n \lfloor \log n \rfloor$, thence the corollary.

It follows from Corollary 5.2 that the message complexity of Algorithm $S_{\text{Elect \_Leader \_C}}$ is $O(n \log n)$. Also, because a capture message carries a node’s identification, it follows that the algorithm’s bit complexity is $O(n \log^2 n)$. This synchronous algorithm has a better message complexity than the one we devised initially (which had $O(n^2)$ message complexity), but this comes at the cost of an increase in time complexity from $O(1)$ to $O(\log n)$. What supports the improved message complexity is the technique of comparing a candidate’s identification to those of its neighbors in increasingly large groups, so that the number of candidates is guaranteed to decrease steadily from an even pulse to another (cf. Theorem 5.1). When we consider the design of an asynchronous counterpart to Algorithm $S_{\text{Elect \_Leader \_C}}$, the use of such a technique has to undergo a few modifications, especially because a node cannot in the asynchronous model consider a group of candidate identifications simultaneously as it did in the synchronous model and reply positively to at
most one of them. It appears, then, that in the asynchronous model a candidate must attempt to capture one node at a time. However, in order to still be able to benefit from the advantages of capturing nodes in groups of increasing sizes, in the asynchronous algorithm identifications are no longer used as a basis of comparison, but rather only to break ties. Comparisons are instead based on the "level" of each competing candidate, which is the number of groups of nodes a candidate has so far succeeded in capturing. This amounts to simulating the technique employed in the synchronous case, but at the expense of a greater time complexity. As we will see, the resulting algorithm, called Algorithm \texttt{A\_Elect\_Leader\_C}, has time complexity \(O(n)\) but its message complexity remains as in the synchronous case, that is, \(O(n \log n)\).

In order to ensure the correctness of this approach, in the sense that no two candidates must ever be allowed to concurrently remain candidates based on having captured a same node, a candidate must only consider a node as having been captured when (and if) that node's current owner ceases being a candidate. The overall approach is then the following. A candidate attempts to capture nodes one at a time. Its level is at all times given by the number of groups it has succeeded in capturing, in the same sense as in Algorithm \texttt{S\_Elect\_Leader\_C}, that is, groups of sizes 1, 2, 4, and so on. If for a candidate \(n\), we let \texttt{level} denote its level and \texttt{owns}, the number of nodes it has captured, then clearly

\[
\texttt{level}(n) = \lceil \log \texttt{owns}(n) \rceil.
\]

In order to capture a node \(n_i\), \(n\) sends it a message \texttt{capture}(\texttt{level}, \texttt{id}). Upon receiving this message, \(n\) checks whether

\[
(\texttt{level}, \texttt{owner.id}) < (\texttt{level}, \texttt{id})
\]

(this comparison is done lexicographically, that is, first the levels are compared and only if they are the same are the identifications compared). If the comparison fails, then \(n\) sends \(n_i\) a \texttt{nack} message, and upon receiving it \(n_i\) ceases being a candidate (if it still is). If, on the other hand, the comparison succeeds, then \texttt{level} is updated to \texttt{level}. In addition, if \(n\) is a candidate, then it ceases being so and \(n\) becomes its owner. Also, \(n\) sends \(n\) an \texttt{ack}, upon receipt of which \(n\) proceeds with its node capturing. If \(n_j\) is not a candidate, then \(n\) is marked as \(n_j\)'s prospective owner. Before \(n\) becomes \(n_j\)'s owner, however, \(n\) has to ensure that \(n_j\)'s current owner ceases being a candidate. To this end, \(n\) sends \(n_j\) a message \texttt{check}(\texttt{k}) (assuming that \texttt{owner.id}(j) = \texttt{id}_n), and upon receiving this message, \(n\), if it still is a candidate, sends a message \texttt{eliminate}(\texttt{level}, \texttt{id}) to \(n_j\). At \(n_k\), the comparison

\[
(\texttt{level}, \texttt{id}_n) < (\texttt{level}, \texttt{id}_k)
\]

is performed and results in one of the following two outcomes. If the comparison fails, then \(n_k\) sends \(n_k\) a \texttt{nack}, thereby causing \(n_k\) not to be a candidate any longer (if it still is). If the comparison succeeds, thereby causing \(n_k\) to cease being a candidate, or if \(n_k\) was no longer a candidate upon receiving the \texttt{eliminate} message, then an \texttt{eliminated} message is sent by \(n_k\) to \(n\), where it causes \(n\) if still a candidate, to try to capture \(n\) once again by sending it another \texttt{capture} message. If this message, upon arriving at \(n\), finds that \(n\) still is \(n_k\)'s prospective owner, then \(n\) becomes \(n_k\)'s new owner and an \texttt{ack} is sent back to \(n_k\). Otherwise, a \texttt{nack} is sent. Upon receipt of one or the other message, \(n\) resumes its captures or ceases being a candidate, respectively. Notice that, throughout this entire process, \(n_k\) has not yet been captured by \(n\), but merely ceased being a candidate.

The variables \texttt{level} and \texttt{owns}, both initially equal to zero, are used by \(n_k\) in Algorithm \texttt{A\_Elect\_Leader\_C} in addition to those already used by Algorithm \texttt{S\_Elect\_Leader\_C}. Node \(n\) employs two other variables, both initialized to \texttt{nil}, to indicate \(n_j\)'s prospective owner and the node it is currently attempting to capture. These are, respectively, \texttt{p\_owner.id}, and \texttt{p\_owned.id}.

\section*{Algorithm \texttt{A\_Elect\_Leader\_C}}

\begin{algorithm}
\begin{itemize}
    \item \texttt{Variables:}
    \begin{itemize}
        \item \texttt{candidate} = \texttt{false};
        \item \texttt{tried}_i^j = \texttt{false} for all \(n_j \in \texttt{Neig}_i\):
        \begin{itemize}
            \item \texttt{owner.id} = \texttt{nil};
            \item \texttt{level} = 0;
            \item \texttt{owns} = 0;
        \end{itemize}
    \end{itemize}
\end{itemize}
\end{algorithm}
\[ p\_owner\_id = \text{nil}; \]
\[ p\_owned\_id = \text{nil}. \]

Listing 5.4

Input:
\[ msg = \text{nil}. \]
Action if \( n \in N \):
\[ \text{candidate} := \text{true}; \]
\[ \text{owner} := id; \]
Let \( n \) be a node in \( \text{Neig} \);
\[ \text{tried}^i := \text{true}; \]
Send \( \text{capture}(\text{level}, id) \) to \( n \).

Listing 5.5

Input:
\[ msg = \text{capture}(\text{level}, id) \text{ such that } \text{origin}(msg) = (n, n). \]
Action:
if \( p\_owner\_id = id \) then
\begin{align*}
& \text{begin} \\
& \quad \text{owner} := id; \\
& \quad \text{Send ack to } n \end{align*}
\end{align*}
else
\begin{align*}
& \text{if } (\text{level}, \text{owner} id) < (\text{level}, id) \text{ then} \\
& \quad \text{begin} \\
& \quad \quad \text{level} := \text{level}; \\
& \quad \quad \text{if } \text{candidate} \text{ then} \\
& \quad \quad \quad \text{begin} \\
& \quad \quad \quad \quad \text{candidate} := \text{false}; \\
& \quad \quad \quad \quad \text{owner} := id; \\
& \quad \quad \quad \quad \text{Send ack to } n \end{align*}
\begin{align*}
& \quad \end{align*}
else
\begin{align*}
& \quad \text{begin} \\
& \quad \quad \text{p} \_\text{owner} \_\text{id} := id; \\
& \quad \quad \text{Let } n \in \text{Neig} \text{ be such that } \text{owner} \_\text{id} = \text{id}; \\
& \quad \quad \text{Send check} (k) \text{ to } n \end{align*}
\begin{align*}
& \quad \end{align*}
else
\begin{align*}
& \quad \text{Send nack to } n. \end{align*}
\begin{align*}
& \end{align*}

Listing 5.6
Input:
\[ \text{msg}_i = \text{nack}. \]

**Action:**
\[ \text{if candidate, then} \]
\[ \quad \text{candidate} := \text{false}. \]

Listing 5.7

Input:
\[ \text{msg}_i = \text{check}(j). \]

**Action:**
\[ \text{if candidate, then} \]
\[ \quad \text{Send eliminate (level, id) to } n_j. \]

Listing 5.8

Input:
\[ \text{msg}_i = \text{eliminate (level, id)} \text{ such that } \text{origin}(\text{msg}) = (n_i, n_j). \]

**Action:**
\[ \text{if not candidate, then} \]
\[ \quad \text{Send eliminated to } n_j \]
\[ \text{else} \]
\[ \quad \text{if (level, id) < (level, id)} \text{ then} \]
\[ \quad \quad \text{begin} \]
\[ \quad \quad \quad \text{candidate} := \text{false}; \]
\[ \quad \quad \quad \text{Send eliminated to } n_j \]
\[ \quad \quad \end{\text{begin} \]
\[ \text{else} \]
\[ \quad \quad \text{Send nack to } n_j. \]

Listing 5.9

Input:
\[ \text{msg}_i = \text{eliminated}. \]

**Action:**
\[ \text{if candidate, then} \]
\[ \quad \text{begin} \]
\[ \quad \quad \text{Let } n_j \in \text{Neig be such that } p\_\text{owned}_i = \text{id}; \]
\[ \quad \quad \text{Send capture (level, id) to } n_j \]
\[ \quad \end{\text{begin} \]

Listing 5.10

Input:
\begin{verbatim}
msg_i = ack.
Action: 
\text{owns}_i := \text{owns}_i + 1;
\text{level}_i := \lfloor \log(\text{owns}_i + 1) \rfloor;

Let \( S \cup \text{Neig}_i \) be such that \( \text{tried}_i = \text{false} \) for all \( n_i \in S \);

if \( S \neq \emptyset \) then
begin
  Let \( n_i \) be a node in \( S \);
  \text{tried}_i := \text{true};
  \text{p-owned-id}_i := \text{id};
  Send \text{capture} (\text{level}_i, \text{id}_i) to \( n_i \);
end.
\end{verbatim}

In Algorithm A_Elect_Leader_C, \((5.4)\) through \((5.10)\) implement the guidelines we gave to employ the technique of Algorithm S_Elect_Leader_C in an asynchronous setting. It should be noted that a candidate \( n_i \) becomes a leader when \( S = \emptyset \) in \((5.10)\). At this time, it must by \((5.5)\) be the owner of all nodes and its level equal to \(|\log n|\). Moreover, by \((5.4)\) through \((5.9)\) a candidate may only be the owner of a node if that node's previous owner is no longer a candidate, which leads us to the following counterpart of Theorem 5.1.

**Theorem 5.3.**

For \( 1 \leq k \leq |\log n| \), the maximum number of candidates of level \( k \) in any global state in an execution of Algorithm A_Elect_Leader_C is \(|n/(2^k - 1)|\).

**Proof:** By the definition of level, a candidate \( n_i \) at level \( k \) must have captured at least \( 2^k - 1 \) of its neighbors, inasmuch as

\[ k = \lfloor \log(\text{owns}_i + 1) \rfloor \leq \log(\text{owns}_i + 1). \]

The theorem then follows from the fact that no two candidates can be owners of a same node in any global state.

**Corollary 5.4.**

Algorithm A_Elect_Leader_C involves at most \( 2n|\log n| + n \) attempts at capturing a node by a candidate.

**Proof:** Before reaching level 1, by \((5.4)\) a candidate attempts to capture exactly one node. For \( 1 \leq k \leq |\log n| \), while at level \( k \) a candidate attempts to capture at most \( 2^k \) nodes. By Theorem 5.3, the total number of node captures the algorithm involves is then

\[ \sum_{k=1}^{\lfloor \log n \rfloor} \frac{n}{2^k - 1} \leq 2n \sum_{k=1}^{\lfloor \log n \rfloor} \frac{n}{2^k - 1^k} = 2n |\log n| + n. \]

Each node capture by a candidate involves at most six messages (one \text{capture}, one \text{check}, one \text{eliminate}, one \text{eliminated}, one more \text{capture}, and one \text{ack}). By Corollary 5.4, the message complexity of Algorithm A_Elect_Leader_C is then \( O(n \log n) \), and because the lengthiest messages (\text{capture} and \text{eliminate} messages) are \(|\log |\log n|| + |\log n| \) bits long, the algorithm's bit complexity is \( O(n \log^2 n) \). In order to check that the time complexity of Algorithm A_Elect_Leader_C is indeed \( O(n) \), it suffices to note that candidates capture nodes independently of one another, in the sense that no candidate depends on another.
candidate's messages to capture nodes (only to cease being a candidate), and that candidates attempt to capture nodes one at a time.

5.2 Distributed snapshots
The second fundamental technique that we discuss in this chapter is a technique for recording global states during the execution of an asynchronous algorithm. While the concept of a global state, as introduced in Section 3.1, is of fundamental importance by itself, the ability to record a global state over which some of the algorithm's global properties can be analyzed is no less attractive. In the context of this book, areas in which algorithms for global state recording are especially relevant include the treatment of stable properties (discussed in Chapter 6) and the handling of some issues related to timing during a distributed simulation (our subject in Chapter 10).

The bulk of our discussion on global state recording is presented in Section 5.2.1, where we present a distributed algorithm to record a global state and leave the recorded information spread throughout the nodes of $G$. However, there are special cases, chiefly within the area of distributed simulation, for which the recording of global states with some specific properties is desirable. In these cases, it seems that the use of a leader to perform the global state recording in a centralized fashion is considerably more efficient. We discuss such a centralized approach in Section 5.2.2, aiming at their use in Chapter 10.

Throughout all of Section 5.2, $G$ is taken to be a directed graph, so that the states of edges can be referred to without explicit mention to a particular direction. The extension to the undirected case is immediate, as usual.

5.2.1 An algorithm
In the case of synchronous algorithms, the recording of a global state can be achieved rather simply. At each pulse $s \geq 0$, the states of all nodes and the messages that were sent at pulse $s - 1$ (if $s > 0$), which by assumption must already have arrived at their destinations, constitute a global state. Without further communication, such a global state can be stored in $G$ distributedly, so that a node stores its own state and the state of all edges on which it receives messages.

Clearly, though, nothing like this simple approach can be employed in the asynchronous case, owing to the total absence of global timing. However, with the aid of communication among the nodes in addition to that pertaining to the computation whose global state we wish to record, the task can also be performed for asynchronous algorithms. The algorithm that we discuss next is surprisingly simple given the apparent intricacy of the task, and yields a global state that can be found at the end of the algorithm stored in a distributed fashion throughout $G$, in much the same way as in the synchronous case we just discussed.

Before we introduce the algorithm for global state recording, it should be noted that, conceptually, we are dealing with two distributed computations. One of them, which we can refer to as the substrate, is the computation whose global properties one wishes to study, and then the global state one is seeking to record is a global state of the substrate. It is then to the substrate that the set of events $\Xi$ introduced in Section 3.1 refers. The other distributed computation is an execution of the algorithm for global state recording, which we henceforth call Algorithm $A_{Record\ Global\ State}$. Both computations run on $G$, so each node is responsible for executing its share of the substrate and of Algorithm $A_{Record\ Global\ State}$. The two computations are, however, totally independent of each other as far as causality relationships are concerned. Our only assumption about their interaction is that Algorithm $A_{Record\ Global\ State}$ is capable of "peeking" at the substrate's variables and messages with the purpose of recording a global state. Note that a node participates in both computations in such a way that, when the substrate is being executed by a node, Algorithm $A_{Record\ Global\ State}$ is suspended, and conversely. This is immaterial from the standpoint of either computation, though. Having been designed to operate in the asynchronous model, the suspension of one to execute the other only adds to the asynchronism already present. Recording a global state during an execution of the substrate is essentially a means of "freezing" that execution in a snapshot (thence this alternative denomination for a global state) to analyze the states of all nodes and edges without actually having to halt the substrate.
This view of the computation at a node as actually comprising the node’s participation in two different distributed algorithms is the view that we adopt in this section. What this amounts to when specifying the actions of Algorithm A_Record_Global_State is that there has to exist an action to handle the receipt of messages of the substrate, although in none of the algorithm’s actions does one such message get sent. Alternatively, we might have viewed both computations as constituting the execution of a single algorithm, in which case the technique for recording global states would appear truly as a building block. When arguing formally about the recorded global state, however, we would have to be careful to discriminate events associated with the substrate from those associated with the additional communication employed by the recording algorithm, as it is to the former that the recorded global state relates.

The following is an outline of how Algorithm A_Record_Global_State functions. A node is responsible for recording the substrate’s local state and the states of all edges directed toward itself. If all nodes carry their recording tasks to their ends, then the resulting overall recording is a system state, as introduced in Section 3.1, because a local state has been recorded for each node and a set of messages for each edge. The algorithm progresses through the exchange between neighbor nodes of a special message called marker. A node \( n_i \in N_0 \) initiates its participation in Algorithm A_Record_Global_State by recording the local state of the substrate, in the terminology of Section 3.1, and then sending marker on all edges that are directed away from it, without however allowing the substrate to send any messages in the meantime (i.e., after the recording of the local state and before the sending of marker). In practice, this can be achieved by "disabling interrupts" so that the node will not switch to execute the other computation while this is undesired. All other nodes behave likewise upon receiving marker for the first time. Every message of the substrate received at \( n \) from a neighbor \( n_j \) after \( n_i \) has received the first marker (and consequently recorded a local state) and before \( n_j \) receives marker from \( n_i \) is added to the set of messages representing \( \Phi_i \), which is the state of edge \( (n_i \rightarrow n_j) \) (cf. Section 3.1 for the appropriate terminology). The state of the edge on which marker was first received is then recorded as the empty set, so the system state recorded by Algorithm A_Record_Global_State can be regarded as containing a forest of empty edges, each of whose trees spanning exactly one node in \( N_0 \). The recording is completed at a node when marker has been received on all edges directed toward that node.

It is instructive at this point to notice the very close resemblance of the algorithm we just outlined to Algorithm A_PI, introduced in Section 4.1.1 for the propagation of information on G. While that algorithm was given for an undirected G, Algorithm A_Record_Global_State can be easily recognized as a variation of Algorithm PI to propagate marker messages by flooding when G is a directed graph. Of course, the question of whether every node in G does ever receive a copy of marker in the directed case arises, because there may exist nodes to which no directed path from a node in \( N_0 \) exists. One situation in which this can be guaranteed is, for example, the case of a strongly connected G, in which a directed path exists from every node to every other node.

Even before describing and analyzing Algorithm A_Record_Global_State more thoroughly, we are then in position to assess its complexities. Because every edge carries at most one copy of marker, the algorithm’s message complexity is clearly \( O(m) \). The algorithm’s time complexity, on the other hand, depends only on how long it takes a marker to reach a node that is not in \( N_0 \), and this is clearly \( O(n) \) time.

In the description of Algorithm A_Record_Global_State we give next, \( sub \_msg \) is used to generically denote a message of the substrate. A node \( n_i \) maintains a variable to store the substrate’s local state at \( n_i \), and for each neighbor \( n_j \in I \_Neig \), a variable to store the state of edge \( (n_i \rightarrow n_j) \). These variables are, respectively, node_state and edge_state, initialized respectively to \( \text{nil} \) and \( \emptyset \). In addition, a variable recorded, (initially equal to false) indicates whether the substrate’s local state has already been recorded, and a variable received, for each \( n_j \in I \_Neig \), (initialized to false as well) indicates whether marker...
has been received from \( n_i \). Clearly, for \( n_i \notin N_0 \), \( recorded_i = \text{true} \) if and only if \( received^j_i = \text{true} \) for some \( n_j \in I_{\text{Neig}} \).

**Algorithm A_Record_Global_State:**

**Variables:**
- \( node\_state_i = \text{nil} \);
- \( edge\_state^j_i = \emptyset \) for all \( n_j \in I_{\text{Neig}} \);
- \( recorded_i = \text{false} \);
- \( received^j_i = \text{false} \) for all \( n_j \in I_{\text{Neig}} \).

**Listing 5.11**

**Input:**

- \( msg_i = \text{nil} \).

**Action** if \( n_i \in N_0 \):
- \( node\_state_i := \sigma_i \);
- \( recorded_i := \text{true} \);
- Send marker to all \( n_j \in O_{\text{Neig}} \).

**Listing 5.12**

**Input:**

- \( msg_i = \text{marker} \) such that \( \text{origin}(msg_i) = (n_j \rightarrow n_i) \).

**Action:**
- \( received^j_i := \text{true} \);
- if not \( recorded_i \) then
  - begin
    - \( node\_state_i := \sigma_i \);
    - \( recorded_i := \text{true} \);
    - Send marker to all \( n_k \in O_{\text{Neig}} \);
  - end.

**Listing 5.13**

**Input:**

- \( msg_i = \text{sub}_msg \) such that \( \text{origin}(msg_i) = (n_j \rightarrow n_i) \).

**Action:**
- if \( recorded_i \) then
  - if not \( received^j_i \) then
There are two important observations to be made concerning Algorithm $A\_Record\_Global\_State$. The first observation is that the assumed atomicity of actions in Algorithm $A\_Template$ (cf. Section 2.1) suffices to prevent a node from executing the substrate computation, possibly with the sending of messages, between the recording of the local state and the sending of marker's in (5.11) and (5.12). The second observation concerns (5.13) and the sub_msg messages that trigger this action. Because such messages are in fact messages of the substrate, the actions that they trigger do not really belong in a presentation of Algorithm $A\_Record\_Global\_State$ based on Algorithm $A\_Template$. In fact, we have made no provisions whatsoever to denote an algorithm's peeking at some other algorithm's messages, so that our notation in the description of Algorithm $A\_Record\_Global\_State$ is abusive. One of the problems caused by this abuse of notation is that Algorithm $A\_Record\_Global\_State$ does not seem to terminate as long as there are $sub\_msg$'s in transit on $G$'s edges, while clearly the algorithm is to terminate as soon as every node $n$ has received as many marker's as there are in $In$ (cf. Section 6.2). As long as these issues are clearly understood, however, our slightly licentious use of the notation should not be troublesome. We keep the improper notation for simplicity (here and in other occasions, as in Section 9.3.3), although it appears that adapting Algorithm $A\_Template$ to properly contemplate such an interaction between two distributed algorithms is a simple matter (cf. Exercise 5).

Theorem 5.5 states two sufficient conditions for the system state that Algorithm $A\_Record\_Global\_State$ records to be a global state.

**Theorem 5.5.**
If $G$ is strongly connected and all of its edges are FIFO, then the system state that Algorithm $A\_Record\_Global\_State$ records is a global state.

**Proof:** The fact that $G$ is strongly connected implies that every node $n$ receives marker exactly once on every edge in $In$, by (5.11) and (5.12). Recalling that $Ξ$ is the set of events related to the substrate only, let $(Ξ₁, Ξ₂)$ be a partition of $Ξ$ such that $ξ \in Ξ₁$ if and only if $ξ$ occurred before the local state of the node at which it occurred was recorded. In addition, referring back to the notation introduced in Section 3.1, let $<^+$ be any total order of the events in $Ξ$ consistent with $<^+$, and consider two consecutive events $ξ₂ \in Ξ₂$ and $ξ₁ \in Ξ₁$ in $<^+$. By the definition of $Ξ₁$ and of $Ξ₂$, it is clear that $ξ₂$ did not happen at the same node as $ξ₁$, and before the occurrence of $ξ₁$. Now consider a scenario in which a sequence of events follow $ξ₂$ at the node at which it happened, and then a message is sent in connection with the last event in this sequence, which in turn eventually causes the sending of another message by its destination node, and then the eventual sending of another message by the destination node of this second message, and so on, and then the arrival of the last message causes a sequence of events to happen at its destination node culminating with the occurrence of $ξ₁$. By (5.11) and (5.12), and by the definition of $Ξ₁$ and $Ξ₂$, the node at which $ξ₂$ happened must have sent marker's before $ξ₂$ happened. Likewise, the node at which $ξ₁$ happened must not have received any marker before $ξ₁$ happened. Clearly, these two requirements are inconsistent with the scenario we just described, as the edges are all FIFO, and the sequence of messages alluded to in the description of the scenario would then have to have been overrun by a marker. In summary, $(ξ₂, ξ₁) \not<^+$, so the total order $<^+$ can be altered by substituting $(ξ₁, ξ₂)$ for $(ξ₂, ξ₁)$ in it, and yet remain consistent with $<^+$. 

$$edge\_state^t_i := edge\_state^t_i \cup \{msg\}.$$
Clearly, it takes no more than $|\Xi_1||\Xi_2|$ such substitutions to obtain a total order in which at most one pair $(\xi_1, \xi_2)$ of consecutive events exists such that $\xi_1 \in \Xi_1$ and $\xi_2 \in \Xi_2$. The events in all other pairs of consecutive events are in this total order both in $\xi_1$, or in $\xi_2$. By (5.11) through (5.13), and by the definition of $\xi_1$ and $\xi_2$, this distinguished pair of consecutive events is such that $\text{system\_state}(\xi_1, \xi_2)$ is precisely the system state recorded by Algorithm A_Record_Global_State, which is then a global state, by our first definition of global states in Section 3.1.

Before we finalize this section, there are a couple of important observations to be made regarding Algorithm A_Record_Global_State. The first observation is that, as we mentioned previously, the global state that the algorithm records is stored in a distributed fashion among $G$'s nodes. Often the recorded global state can be used without having to be concentrated on a single node for analysis (cf. Section 6.3.2 for an example), but equally as frequently it must first be concentrated on a leader, which then works on the global state in a centralized manner.

The second observation is that the global state that the algorithm records is in principle any global state, in the sense that no control is provided to make "choices" regarding desirable characteristics of the resulting global state. While this is fine for many applications (as for example the detection of the stable properties we treat in Chapter 6), for others it does matter which global state is used, and then a centralized approach may be advisable. We elaborate on this a little more in Section 5.2.2.

We end the discussion in this section by returning to the issue of knowledge in distributed computations, treated in Section 2.3, in order to illustrate one of the concepts introduced in that section. Specifically, let $P$ be any sentence related to a global state that has been recorded by Algorithm A_Record_Global_State. Because of the distributed fashion in which this global state is stored after it is recorded, $P$ is clearly implicit knowledge that the members of $N$ have, that is, $I_N P$.

5.2.2 Some centralized alternatives

In this section, we briefly comment on two centralized alternatives to the recording of global states, drawing mainly on motivations to be found in Chapter 10. As we mentioned in the previous section, these centralized alternatives are a solution when the global state that one seeks to record cannot be just any global state, but instead must be a global state with certain specific characteristics.

The first case that we examine is that of a computation for which it is known that the system state in which every node $n_i$ is in the $k$th local state in the sequence $\sigma_i$ (cf. Section 3.1) for some $k \geq 1$ is a global state. This is the case, for example, of the algorithms we discuss in Section 10.2, where for $k > 1$ a function of the $k$th such global state has to be compared with the result of applying the same function to the $k - 1$st such global state, regardless of whatever messages may be in transit on the edges. It turns out that computing such a function of a global state is a trivial task if done by a single node, and because the edge states do not matter, the natural choice is for every node $n_i$ to report every new local state to a leader, which then performs the necessary function evaluation and global state comparison whenever a new global state is completed with information received from the other nodes.

The second case is motivated by the needs of Sections 10.3.2 and 10.6, where in the global states to be recorded every edge state must be the empty set. Again, a centralized approach is preferable because it renders the task of checking for empty edges quite simple. The approach is then the following. Whenever a node $n_i$ is in a local state with which it may participate in a global state of interest (or simply periodically), it sends the leader this local state, together with the numbers of messages it has so far received on each edge in $\text{In}_i$ and sent on each edge in $\text{Out}_i$. Whenever the information the leader receives from all nodes is such that the number of messages sent on each edge is equal to the number of messages received on that edge, the corresponding system state is surely a global state, because every system state in which all edges are empty is a global state (cf. Exercise 7).
5.3 Network synchronization

This section is dedicated to the third major design technique to be discussed in this chapter, namely network synchronization. As we have anticipated in various occasions, especially in Sections 2.1 and 3.3, a synchronous algorithm can be turned into an asynchronous algorithm at the expense of additional message and time complexities, so that the lack of a global time basis and of bounds on delays for message delivery can be dealt with, and the resulting algorithm can be guaranteed to function as in the synchronous model.

As we remarked in Section 3.3 when presenting Algorithm S-to-A-Template, essentially the technique of network synchronization amounts to determining that a node \( n \) that has been executing the action corresponding to pulse \( s \geq 0 \) of the synchronous algorithm is ready to proceed to pulse \( s + 1 \) under the asynchronous model. In Algorithm S-to-A-Template, such a decision is embodied in a Boolean function \( \text{DONE}(s) \), where \( s \) is a variable that indicates the interval \( n \) is currently involved with.

Our approach in this section is to take \( G \) to be an undirected graph, and then consider a generic synchronous algorithm, call it Algorithm \( S_{-}\text{Alg} \), written in accordance with Algorithm \( S_{-}\text{Template} \) of Section 2.1. The asynchronous algorithm resulting from translating Algorithm \( S_{-}\text{Alg} \) into the asynchronous model will be called Algorithm \( A_{-}\text{Alg} \), where \( \text{Sync} \) indicates the particular technique, or synchronizer, employed in the translation. In essence, Algorithm \( A_{-}\text{Alg} \) follows Algorithm S-to-A-Template.

The essential property that we seek to preserve in translating Algorithm \( S_{-}\text{Alg} \) into Algorithm \( A_{-}\text{Alg} \) is that no node \( n \), proceeds to pulse \( s + 1 \) before all messages sent to it at pulse \( s \) have been delivered and incorporated into MSG(\( s \)) (the reader should recall from Sections 2.1, and 3.3, that this is the set of messages sent to \( n \), at pulse \( s \)). In order to ensure that this property holds for all nodes and at all pulses, we begin by requiring that all messages of Algorithm \( S_{-}\text{Alg} \) be acknowledged. These messages are denoted by \( \text{comp}_\text{msg} \), and the acknowledgements by \( \text{ack} \). A node is said to be safe with respect to pulse \( s \) if and only if it has received an \( \text{ack} \) for every \( \text{comp}_\text{msg} \) it sent at pulse \( s \). In order to guarantee that our essential property holds for \( n \), at pulse \( s \), it then suffices that \( n \) receive information stating that every one of its neighbors is safe with respect to pulse \( s \). The task of a synchronizer is then to convey this information to all nodes concerning all pulses of the synchronous computation.

A synchronizer is then to be understood as an asynchronous algorithm that is repeated at every pulse of Algorithm \( S_{-}\text{Alg} \) in order to convey to all nodes the safety information we have identified as fundamental. Now let Messages(\( \text{Alg} \)) and Time(\( \text{Alg} \)) denote, respectively, the message complexity and the time complexity of a distributed algorithm \( \text{Alg} \) (synchronous or asynchronous). Then Messages(\( \text{Sync} \)) and Time(\( \text{Sync} \)) stand for the message and time complexities, respectively, introduced by Synchronizer \( \text{Sync} \) per pulse of Algorithm \( S_{-}\text{Alg} \) to yield Algorithm \( A_{-}\text{Alg} \). These two quantities constitute the synchronization overhead introduced by Synchronizer \( \text{Sync} \).

Regardless of how Synchronizer \( \text{Sync} \) operates, we can already draw some conclusions regarding the final complexities of Algorithm \( A_{-}\text{Alg} \). Let us, first of all, recognize that the use of the \( \text{ack} \) messages does not add to the message complexity of Algorithm \( S_{-}\text{Alg} \), as exactly one \( \text{ack} \) is sent per \( \text{comp}_\text{msg} \). Considering in addition that Messages(\( \text{Sync} \)) is the message complexity introduced by Synchronizer \( \text{Sync} \) per pulse of the execution of Algorithm \( S_{-}\text{Alg} \), and that there are Time(\( S_{-}\text{Alg} \)) such pulses, we then have

\[
\text{Messages}(A_{-}\text{Alg}(\text{Sync})) = \text{Messages}(S_{-}\text{Alg}) + \text{Time}(S_{-}\text{Alg})\text{Messages}(\text{Sync}) + \text{Messages}(\text{Sync}),
\]

where Messages(\( \text{Sync} \)) is the message complexity, if any, that Synchronizer \( \text{Sync} \) incurs with initialization procedures.

Similarly, as Time(\( \text{Sync} \)) is the time complexity introduced by Synchronizer \( \text{Sync} \) per each of the Time(\( S_{-}\text{Alg} \)) pulses of Algorithm \( S_{-}\text{Alg} \), we have

\[
\text{Time}(A_{-}\text{Alg}(\text{Sync})) = \text{Time}(S_{-}\text{Alg})\text{Time}(\text{Sync}) + \text{Time}(\text{Sync}),
\]

where Time(\( \text{Sync} \)) refers to the time, if any, needed by Synchronizer \( \text{Sync} \) to be initialized. Depending on how Synchronizer \( \text{Sync} \) is designed, the resulting complexities Messages(\( A_{-}\text{Alg}(\text{Sync}) \)) and Time(\( A_{-}\text{Alg}(\text{Sync}) \)) can vary considerably. In Section 5.3.1, we discuss three types of general synchronizers, and in Section 5.3.2, consider some special variations of interest.
5.3.1 General synchronizers
The essential task of a synchronizer is to convey to every node and for every pulse the information that all of the node's neighbors are safe with respect to that pulse. This safety information indicates that the node's neighbors have received an ack for every comp_msg they sent at that pulse, and therefore the node may proceed to the next pulse.

The first synchronizer we present is known as Synchronizer Alpha. The material that we present in Section 5.3.2 comprises variants of this synchronizer. In Synchronizer Alpha, the information that all of a node’s neighbors are safe with respect to pulse $s \geq 0$ is conveyed directly by each of those neighbors by means of a safe$(s)$ message. A node may then proceed to pulse $s + 1$ when it has received a safe$(s)$ from each of its neighbors. Clearly, we have

\[
\text{Messages}(\text{Alpha}) = O(n),
\]

and

\[
\text{Time}(\text{Alpha}) = O(1),
\]
as a safe message is sent between each pair of neighbors in each direction, and causes no effect that propagates farther than one edge away. We also have \(\text{Messages}_0(\text{Alpha}) = \text{Time}_0(\text{Alpha}) = 0\).

Algorithm $A$ _Alg(Alpha) is described next. In this section, we do not assume that edges are FIFO, and for this reason comp_msg's and ack's sent in connection with pulse $s \geq 0$ are sent as comp-msg$(s)$ and ack$(s)$ (cf. Exercise 8). In Algorithm $A$ _Alg(Alpha), node $n$, maintains, in addition to the variables employed by Algorithm $S$-to-$A$ Template, the following others. A variable \(\text{expected}(s)\), initially equal to zero, records for all $s \geq 0$ the number of ack$(s)$'s $n_i$ expects. This variable is assumed to be incremented accordingly whenever $n_i$ sends comp_msg$(s)$'s, although this is part of the "Send one message..." that generically appears in all our templates and then the sending of the messages is not explicitly shown. Node $n_i,

\[
safe^i_j(s)
\]
also maintains a variable $\text{safe}(s)$ for each neighbor $n_i$ and all $s \geq 0$, initially set to true and used to indicate whether a safe$(s)$ has been received from $n_i$.

Despite the simplicity of Synchronizer Alpha, designing the initial actions of Algorithm $A$ _Alg(Alpha) requires that we reason carefully along the following lines. A node in $N_0$ behaves initially just as it would in the synchronous model. A node in that is not in $N_0$, however, although in Algorithm $S$-Alg it might remain idle for any number of pulses, in Algorithm $A$-Alg(Alpha) it must take actions corresponding to every pulse, because otherwise its neighbors would never receive the safe messages that it should send and then not progress in the computation. The way we approach this is by employing an additional message, called startup, which is sent by the nodes in $N_0$ to all of their neighbors when they start computing.

This message, upon reaching a node that is not in $N_0$ for the first time, serves the purpose of "waking" that node up and then gets forwarded by it to all of its neighbors as well. Loosely, this startup message can be thought of as a "safe($\text{-}1$)" message that is propagated in the manner of Algorithm $A$-PI of Section 4.1.1, and is intended to convey to the nodes that are not in $N_0$ the information that they should participate in pulse $s = 0$ too, as well as in all other pulses (although for $s > 0$ this can be taken for granted by the functioning of Synchronizer Alpha). All nodes, including those in $N_0$, only proceed to executing pulse $s = 0$ of the synchronous computation upon receiving a startup from every neighbor. This is controlled by a variable $\text{initiated}_i$, initially set to true, maintained by $n_i$ for every neighbor $n_i$ to indicate whether a startup has been received from $n_i$. An additional variable, initiated, initially set to false as well, indicates whether $n_i \in N_0$.

**Algorithm $A$ _Alg(Alpha):**

<table>
<thead>
<tr>
<th>Variables</th>
</tr>
</thead>
<tbody>
<tr>
<td>$s_i = 0$;</td>
</tr>
</tbody>
</table>

\[
\text{MSG}(s) = \emptyset \quad \text{for all } s \geq 0;
\]
\[ \text{initiated}_i = \text{false}; \]
\[ \text{go}^j_i = \text{false} \text{ for all } n_j \in \text{Neig}_i; \]
\[ \text{expected}(s) = 0 \text{ for all } s \geq 0; \]
\[ \text{safe}^j_i(s) = \text{false} \text{ for all } n_j \in \text{Neig}_i \text{ and all } s \geq 0. \]

**Listing 5.14**

**Input:**
\[ \text{msg}_i = \text{nil}. \]
**Action if \( n_i \in \text{Neig}_i: \)**
\[ \text{initiated}_i := \text{true}; \]
Send startup to all \( n_j \in \text{Neig}_i. \]

**Listing 5.15**

**Input:**
\[ \text{msg}_i = \text{startup such that origin}(\text{msg}_i) = (n_i, n_j). \]
**Action:**
\[ \text{if not } \text{initiated}_i \text{ then} \]
\[ \text{begin} \]
\[ \text{initiated}_i := \text{true}; \]
\[ \text{Send startup to all } n_k \in \text{Neig}_i; \]
\[ \text{end}; \]
\[ \text{go}^j_i := \text{true}; \]
\[ \text{if } \text{go}^j_i \text{ for all } n_j \in \text{Neig}_i \text{ then} \]
\[ \text{begin} \]
\[ \text{Do some computation;} \]
\[ \text{Send one } \text{comp_msg}(s) \text{ on each edge of a (possibly empty)} \]
\[ \text{subset of } \text{Inc} \text{;} \]
\[ \text{if } \text{expected}(s) = 0 \text{ then} \]
\[ \text{Send } \text{safe}(s) \text{ to all } n_k \in \text{Neig}_i; \]
\[ \text{end}. \]

**Listing 5.16**

**Input:**
\[ \text{msg}_i = \text{comp_msg}(s) \text{ such that origin}(\text{msg}_i) = (n_i, n_j). \]
**Action:**
\[ \text{MSG}(s + 1) := \text{MSG}(s + 1) \cup \{ \text{msg}_i \}; \]
Send \text{ack}(s) to \( n_i. \)
Listing 5.17

Input:
\( msg_i = \text{ack}(s) \).

Action:
\[
\text{expected}(s) := \text{expected}(s) - 1; \\
\text{if } \text{expected}(s) = 0 \text{ then} \\
\quad \text{Send safe}(s) \text{ to all } n_j \in \text{Neig}_i.
\]

Listing 5.18

Input:
\( msg_i = \text{safe}(s) \) such that \( \text{origin}(msg_i) = (n_i, n_j) \).

Action:
\[
\text{safe}_i^j(s) := \text{true} \\
\text{if } \text{safe}_i^k(s) \text{ for all } n_k \in \text{Neig}_i \text{ then} \\
\quad \text{begin} \\
\quad \quad s_i := s_i + 1; \\
\quad \quad \text{Do some computation;} \\
\quad \quad \text{Send one comp_msg}(s_i) \text{ on each edge of a (possibly empty) subset of Inc}_i; \\
\quad \quad \text{if } \text{expected}(s_i) = 0 \text{ then} \\
\quad \quad \quad \text{Send safe}(s_i) \text{ to all } n_k \in \text{Neig}_i; \\
\quad \quad \text{end.}
\]

As we indicated earlier, Algorithm \( A_{\text{Alg}}(\text{Alpha}) \) can be viewed as a specialization of Algorithm S-to-A_Template when the synchronization technique is Synchronizer \( \text{Alpha} \). Indeed, the reader may without any difficulty check that (5.14) and (5.15) essentially do the job of (3.3), although the former involve nodes that are not in \( N_0 \) while the latter does not. Similarly, (5.16) through (5.18) offer a detailed view of (3.4) under the rules of Synchronizer \( \text{Alpha} \). In particular \( \text{DONE}_i(s) \) returns a \textbf{true} value in (3.4) if and only if \( \text{safe}_i^j(s) = \text{true} \) for all \( n_j \in \text{Neig}_i \) in (5.18).

Synchronizer \( \text{Alpha} \) is only one of the possibilities. For generic synchronous computations like Algorithm S_Alg, there are two other synchronizers of interest. The first one is called Synchronizer \( \text{Beta} \), and requires for its operation a spanning tree already established on \( G \), so the initial complexities \( \text{Messages}_0(\text{Beta}) \) and \( \text{Time}_0(\text{Beta}) \) are no longer equal to zero, but depend instead on the distributed algorithm used to generate the tree (cf. Section 7.1.2). These complexities must also account for the election of a leader, which, as we mentioned in Section 5.1, may be carried out rather closely to the construction of the spanning tree (cf. Section 7.1.1, and Section 7.1.2 as well).

The function of the leader in Synchronizer \( \text{Beta} \) is to gather from all other nodes the safety information needed to proceed to further pulses, and then broadcast this information to all of them. The specifics of this procedure are the following. When a node that is not the leader
becomes safe with respect to a certain pulse and has received a safe message from all but one of its neighbors on the tree, it then sends a safe message to the single neighbor from which it did not receive a safe (the tree edge connecting to this neighbor leads towards the leader). The leader, upon receiving safe messages on all the tree edges that are incident to it, and being itself safe with respect to that pulse, broadcasts a message on the tree indicating that the computation of a new pulse may be undertaken. This message may be a safe message as well, and then the rule for a node to proceed to another pulse is to do it after having received a safe message on all tree edges incident to it. Once the leader has been elected and the spanning tree built, the asynchronous algorithm that results from applying Synchronizer Beta to Algorithm S_Alg, Algorithm A-Alg(Beta), is initiated as follows. The leader broadcasts on the tree that all nodes may begin the computation of pulse $s = 0$.

Clearly, the messages that Synchronizer Beta introduces traverse only tree edges, so we have

$$\text{Messages}(\text{Beta}) = O(n)$$

and

$$\text{Time}(\text{Beta}) = O(n).$$

For generic computations, Synchronizer Beta does better than Synchronizer Alpha in terms of message complexity, whereas the reverse holds in terms of time complexity.

The other synchronizer of interest, called Synchronizer Gamma, arises from a combination of Synchronizers Alpha and Beta. In this combination, nodes are conceptually grouped into clusters. Inside clusters, Synchronizer Gamma operates as Synchronizer Alpha; among clusters, it operates as Synchronizer Beta. The size and disposition of clusters are regulated by a parameter $k$ such that $2 \leq k < n$, and in such a way that Synchronizer Gamma’s complexities are

$$\text{Messages}(\text{Gamma}) = O(\log n \log k)$$

and

$$\text{Time}(\text{Gamma}) = O\left(\frac{n \log n}{\log k}\right).$$

As $k$ varies, Synchronizer Gamma resembles more Synchronizer Alpha or Synchronizer Beta. Once again the costs of initialization $\text{Messages}_0(\text{Gamma})$ and $\text{Time}_0(\text{Gamma})$ are nonzero and depend on the mechanisms utilized. Values that can be attained for these measures are $\text{Messages}_0(\text{Gamma}) = O(kn^2)$ and $\text{Time}_0(\text{Gamma}) = O(n \log n / \log k)$.

Something instructive for the reader to do, having become acquainted with the synchronizers we discussed in this section, is to return to the various synchronous algorithms we have already seen in the book and assess their complexities when each of the three synchronizers is employed. Some of the conclusions to be drawn from this assessment are the following.

First, no synchronizer can beat the complexities of Algorithm A-Compute-f when applied to either Algorithm S-Compute-AND or Algorithm S-Locally-Orient. Secondly, Synchronizer Alpha (the only one whose application to a synchronous leader election algorithm is meaningful), when applied to Algorithm S-Elect-Leader-C, does not yield improvements over the complexity of Algorithm A-Elect-Leader-C.

What these two conclusions indicate is that synchronizers do not necessarily lead to better complexities when compared with asynchronous algorithms that were designed without recourse to synchronization techniques. However, as we remarked at the end of Section 3.4 in the context of establishing a breadth-first numbering on the nodes of a directed graph, historically there have been occasions in which such improvements were obtained.

Incidentally, it may also be an instructive exercise for the reader to verify that the complexities claimed in that occasion for the asynchronous solution obtained from the synchronous one are consistent with the message and time complexities of Synchronizer Gamma as discussed in this section. (Although in this section $G$ is taken to be undirected, no conflict exists when addressing the computation for breadth-first numbering discussed in Section 3.4. The graph is in that case a directed graph, and it is on such a graph that the
synchronous algorithm operates. On the other hand, the synchronizer, and consequently the resulting asynchronous algorithm, operate on the corresponding undirected graph, essentially by being allowed to send synchronization-related messages against the direction of the edges when needed—cf. Exercise 10.) In Chapter 7, when we discuss algorithms to find maximum flows in networks, synchronizers will come to the fore once again in the book.

5.3.2 Important special cases

Of the synchronous algorithms we have seen so far in the book, another that deserves our attention in the light of a synchronizer is Algorithm S-Compute-Distances, introduced in Section 4.3 for the computation of the shortest distances among all pairs of nodes in G. As we claimed in that section, this algorithm yields Algorithm A-Compute-Distances, also introduced in Section 4.3, through the utilization of a synchronizer. Interestingly, and contrary to the intuition we built during our study in the previous section, both algorithms have the same message and time complexities. The reason for this is that the synchronizer employed to transform the synchronous algorithm into the asynchronous one is a particular case of Synchronizer Alpha, as we discuss next.

Let us first, however, examine the following scenario in the context of Synchronizer Alpha. Suppose, for the sake of example, that node \( n_i \) has two neighbors, \( n_j \) and \( n_k \). Suppose further that \( n_j \) has only one neighbor (\( n_i \)) and that \( n_k \) has many neighbors. Consider the situation in which \( n_j \) has just become safe with respect to pulse \( s \geq 0 \) and then sends \( \text{safe}(s) \) to \( n_j \) and \( n_k \). Node \( n_j \) can only proceed to pulse \( s + 1 \) after receiving similar messages from its two neighbors. Suppose that such a message has been received from \( n_j \) but not from \( n_k \) (which depends on many more neighbors other than \( n_j \) to become safe with respect to pulse \( s \)). It is possible at this moment that a \( \text{safe}(s + 1) \) too is received from \( n_j \) before the \( \text{safe}(s) \) from \( n_k \) arrives, at which time \( n_j \) will have received two \( \text{safe} \) messages from \( n_j \) without the respective counterparts from \( n_k \), and will therefore be unable to proceed to pulse \( s + 1 \) immediately. It is simple to see, nevertheless, that \( n_k \) will at this time be unable to proceed to pulse \( s + 2 \), as it now depends on a \( \text{safe}(s + 1) \) from \( n_j \). If we compute the number of \( \text{safe} \) messages \( n_j \) has received since the beginning of the computation from its two neighbors, we will see that the numbers corresponding to \( n_j \) and \( n_k \) differ by no more than two. The particular topological situation we described was meant to help the understanding of this issue, but the maximum difference we just stated is true in general.

This relative "boundedness," when coupled with the assumption that all edges are FIFO, allows various simplifications to be carried out on Algorithm A-Alg(Alpha). For one thing, no message or variable needs to depend explicitly on \( s \) any longer, so that the "per-pulse bit complexity" of Synchronizer Alpha, which we have not introduced formally but clearly might be unbounded, becomes constant. In addition, under the FIFO assumption \( \text{ack} \)'s are no longer needed, and \( n_i \) may send \( \text{safe} \) messages to all of its neighbors immediately upon completion of its computation for the corresponding pulse. Such \( \text{safe} \) messages will certainly be delivered after the \( \text{comp-msg} \)'s sent during that computation, indicating that every such message sent by \( n_i \) during the current pulse has already arrived.

We now present a version of Algorithm A-Alg (Alpha) in which edges are assumed to be FIFO, and, in addition, the following important assumption is made. At each pulse \( s \geq 0 \), node \( n_i \) sends exactly one \( \text{comp-msg} \) to each of its neighbors. These two assumptions allow \( \text{startup}, \text{ack}, \) and \( \text{safe} \) messages to be done away with altogether, and so render the variables \( g_{i,j}^j, \text{expected}^j, \) and \( \text{safe}^j \) useless for all neighbors \( n_i \) of \( n_j \). Also, the sets \( \text{MSG}^j \) (s) for \( s \geq 0 \) are no longer needed; instead, one single set \( \text{MSG}^i \) for use at all pulses suffices, as we see next.

The behavior of \( n_j \) is now considerably simpler, and goes as follows. It starts upon receiving the first \( \text{comp-msg} \) (unless it belongs to \( N_0 \)), and proceeds to the next pulse upon receiving exactly one \( \text{comp-msg} \) from each of its neighbors. However, it is still possible to receive two consecutive \( \text{comp-msg} \)'s from one neighbor without having received any \( \text{comp-msg} \) from another neighbor. This issue is essentially the same we discussed above concerning the reception of multiple \( \text{safe} \) messages from a same neighbor, and some control mechanism has to be adopted. What we need is, for each neighbor, a queue with one single position in
which \textit{comp\_msg}'s received from that neighbor are kept until they can be incorporated into $MSG$. (From our previous discussion, it would seem that two-position queues are needed. However, we can think of $MSG_i$ as containing the queue heads for all of $n_i$'s queues.) We then let $queue^j_i$ denote this queue at $n_i$ for neighbor $n_j$. The new version we present is called Algorithm A\_Schedule\_AS ("AS" for Alpha Synchronization), in allusion to its use in Section 10.2.

**Algorithm A\_Schedule\_AS:**

**Variables:**
- $s_i = 0$
- $MSG_i = \emptyset$
- $initiated_i = \text{false}$
- $queue^j_i = \text{nil}$ for all $n_j \in \text{Neig}_i$.

**Listing 5.19**

**Input:**
$\text{msg}_i = \text{nil}$.

**Action if $n_i \in N_0$:**
- $\text{initiated}_i := \text{true}$;
- Do some computation;
- Send exactly one $\text{comp\_msg}$ on each edge of $Inc_i$.

**Listing 5.20**

**Input:**
$\text{msg}_i = \text{comp\_msg}$ such that $\text{origin}(\text{msg}_i) = (n_i, n_j)$.

**Action:**
- if not $\text{initiated}_i$ then
  - $\text{initiated}_i := \text{true}$;
  - Do some computation;
  - Send exactly one $\text{comp\_msg}$ on each edge of $Inc_i$;
- if there exists $\text{msg}_i \in MSG_i$ such that $\text{origin}(\text{msg}_i) = n_j$ then
  - $queue^j_i := \text{msg}_i$;
- else
  - $MSG_i := MSG_i \cup \{\text{msg}_i\}$;
  - if $|MSG_i| = |\text{Neig}_i|$ then
    - $s_i := s_i + 1$;
    - Do some computation;
    - Send exactly one $\text{comp\_msg}$ on each edge of $Inc_i$;
    - $MSG_i := \emptyset$;
In Algorithm $A_{Schedule\_AS}$, (5.19) and (5.20) reflect the considerable simplification that the assumptions of this section entail with respect to Algorithm $A_{Alg}(\alpha)$. In addition to the elimination of many messages and variables with respect to that algorithm, it should also be noted that, unless $n_i$ employs the value of $s_i$ for its computation at any pulse, this variable too may be eliminated.

When comparing this algorithm with the general template given by Algorithm $S\rightarrow A_{Template}$, one verifies that $DONE(s_i)$ returns true in (3.4) if and only if $|MSG| = |Neig|$ in (5.20), although the dependency on $s_i$ is no longer explicit, as $MSG_i$ is a single set for use at all pulses.

We are now in position to return to the problem of computing shortest distances where we left it in Section 4.3. Clearly, Algorithm $S_{Compute\_Distances}$ complies with the assumption of this section that every node sends exactly one message to every one of its neighbors at all pulses. This, combined with the assumption of FIFO edges, allows a corresponding synchronous algorithm to be obtained along the lines of Algorithm $A_{Schedule\_AS}$. Indeed, it should take little effort to realize that Algorithm $A_{Compute\_Distances}$ is merely an instance of Algorithm $A_{Schedule\_AS}$ (cf. Exercise 11). Because the latter, when viewed as a synchronous algorithm that underwent synchronization, does not contain any synchronization overhead, the complexities of Algorithm $A_{Compute\_Distances}$ are indeed the same as those of Algorithm $S_{Compute\_Distances}$.

5.4 Exercises

1. Show that a leader can only be elected if in $G$ all nodes have distinct identifications.
2. Discuss what happens to Algorithm $S_{Elect\_Leader\_C}$ if the base is no longer $2$, but rather $c$ such that $2 \leq c < n - 1$.
3. Consider the $O(n^2)$-message, $O(1)$-time synchronous algorithm that we discussed in Section 5.1 for leader election on a complete graph, and discuss how it can be adapted to the asynchronous case. Show that the message complexity remains the same, but the time complexity becomes $O(n)$. Compare the resulting algorithm with Algorithm $A_{Elect\_Leader\_C}$.
4. Derive a leader-election algorithm from multiple executions of Algorithm $A_{PI}$.
5. Discuss alternatives to Algorithm $A_{Template}$ that allow the treatment of messages belonging to another computation as well. Show how this affects the way Algorithm $A_{Record\_Global\_State}$ is expressed.
6. Consider a computation in which nodes halt independently of one another, and consider the system state in which all nodes are halted. Is this system state a global state? If it is, is it completely known to the nodes?
7. Show that every system state in which all edges are empty is a global state.
8. Discuss what may happen if the pulse number is omitted from the messages $comp\_msg$ and $ack$ in Algorithm $A_{Alg}(\alpha)$ when edges are not FIFO.
9. In the context of Section 1.5, find the $r(c)$'s for Algorithm $A_{Schedule\_AS}$.
10. Discuss the fundamental alterations synchronizers must undergo when $G$ is a directed graph.
11. Discuss in detail the reasons why Algorithm $A_{Compute\_Distances}$ is an instance of Algorithm $A_{Schedule\_AS}$.
12. Explain how to modify Algorithm $A_{Compute\_Distances}$ so that useless work is avoided after all distances have been determined (instead of keeping running up to the maximum possible distance of $n - 1$).

1. Show that a leader can only be elected if in $G$ all nodes have distinct identifications.
2. Discuss what happens to Algorithm S_Elect_Leader_C if the base is no longer 2, but rather \( c \) such that \( 2 \leq c < n - 1 \).

3. Consider the \( O(n^2) \)-message, \( O(1) \)-time synchronous algorithm that we discussed in Section 5.1 for leader election on a complete graph, and discuss how it can be adapted to the asynchronous case. Show that the message complexity remains the same, but the time complexity becomes \( O(n) \). Compare the resulting algorithm with Algorithm A_Elect_Leader_C.

4. Derive a leader-election algorithm from multiple executions of Algorithm A_PI.

5. Discuss alternatives to Algorithm A_Template that allow the treatment of messages belonging to another computation as well. Show how this affects the way Algorithm A_Record_Global_State is expressed.

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7. Show that every system state in which all edges are empty is a global state.

8. Discuss what may happen if the pulse number is omitted from the messages comp_msg and ack in Algorithm A_Alg(Alpha) when edges are not FIFO.

9. In the context of Section 1.5, find the \( r(c) \)'s for Algorithm A_Schedule_AS.

10. Discuss the fundamental alterations synchronizers must undergo when \( G \) is a directed graph.

11. Discuss in detail the reasons why Algorithm A_Compute_Distances is an instance of Algorithm A_Schedule_AS.

12. Explain how to modify Algorithm A_Compute_Distances so that useless work is avoided after all distances have been determined (instead of keeping running up to the maximum possible distance of \( n - 1 \)).

5.5 Bibliographic notes

The impossibility of electing leaders in the absence of distinct identifications for all nodes is discussed in Angluin (1980). Our treatment in Section 5.1 is based on Afek and Gafni (1991). Many other authors have investigated the problem of electing a leader under various restrictions on \( G \), including synchronous rings (Overmars and Santoro, 1989; Bodlaender and Tel, 1990) and cases in which \( G \) is directed (Afek and Gafni, 1994). Awerbuch (1987) has addressed the problem for generic graphs, and appears to have given the first time-optimal algorithm to solve it—cf. Section 7.4. Additional work on leader election includes the contributions by Peleg (1990), Singh (1992), Tsaan Huang (1993), and Singh and Kurose (1994).

Most of Section 5.2 is based on the seminal work by Chandy and Lamport (1985). Other authors have recently addressed the problem of global state recording in different contexts, as for example Acharya and Badrinath (1992), Alagar and Venkatesan (1994), and Saleh, Ural, and Agarwal (1994). Applications of algorithms for global state recording other than those presented in other chapters can be found in Chaves Filho and Barbosa (1992) and in Choy and Singh (1993), in both cases for scheduling purposes.

A great portion of Section 5.3 is based on the work in which synchronizers were first introduced (Awerbuch, 1985a). Further developments on the theme can be looked up in Awerbuch and Peleg (1990), Shabtay and Segall (1992), Garofalakis, Spirakis, Tampakas, and Rajsbaum (1994), and Rajsbaum and Sidi (1994).
Distributed snapshots and synchronizers are often regarded as essential building blocks for the design of distributed algorithms in general. The reader interested in such a view of the design of distributed algorithms may refer to Gafni (1986), and to additional publications in which techniques with potential to occupy similar positions as building blocks have been introduced (Afek, Awerbuch, and Gafni, 1987; Afek and Ricklin, 1993).

Part 2: Advances and Applications

Stable Properties
Graph Algorithms
Resource Sharing
Program Debugging
Simulation

This second part of the book comprises five additional chapters, each dedicated to a class of problems for which distributed algorithms have been devised. These algorithms constitute advances on the basic algorithms and techniques introduced in the chapters of Part 1, and are in most cases geared toward particular classes of applications.

Chapter 6 contains a study of stable properties from the standpoints of self-stabilization and of stability detection. The investigation of self-stabilizing computations may be ultimately applicable to the recovery from faults, while the detection of stability finds much more immediate applicability, for example in the areas of termination and deadlock detection, both discussed in the chapter. Chapter 7 expands on material seen previously in Part 1 (Chapters 4 and 5) with the study of two graph problems. The first graph problem is that of finding a minimum spanning tree on a graph, and relates directly to the leader election problem, studied in Chapter 5. The second graph problem is that of finding a maximum flow in a graph with a few special features. This problem, like those seen in Chapter 4, are related to problems in the operation of distributed-memory systems.

Chapter 8 is dedicated to the study of distributed algorithms to ensure mutual exclusion in the access to shared resources, while guaranteeing deadlock- and starvation-freedom as well. This problem is studied from two broad perspectives, which in essence can be reduced to the sharing of one single resource or of multiple resources concomitantly. One of the algorithms studied in this chapter provides the basis for part of the discussion in Chapter 10. Techniques for the deterministic re-execution of distributed algorithms in an asynchronous setting, and for detecting breakpoints during executions of such algorithms, are studied in Chapter 9. Both problems constitute essential parts of the process of program debugging, and present difficulties far beyond those encountered in a sequential setting. For the detection of breakpoints, we restrict our attention to a few classes of breakpoints only.

Chapter 10 contains material on the distributed simulation of physical systems, which are models of natural systems occurring in various scientific fields. We present approaches for two broad classes of systems, called the time-stepped and event-driven approaches. Within the latter, we expand on the so-called conservative and optimistic methods. The chapter also contains a brief discussion of methods for systems that do not exactly fall into either of the two classes, as well as a short digression on how the various approaches may be unified.

Chapter 6: Stable Properties

Overview

A stable property is a global property of $G$ that holds for all global states in the future of a global state for which it holds. This chapter is devoted to the study of stable properties from two essentially distinct perspectives. The first perspective is that of ensuring that a stable property is achieved regardless of the initial global state, and the second perspective is that of detecting that a stable property holds for some global state.
We address stable properties as a desired goal from any initial global state in Section 6.1, where we relate such a type of behavior to the issue of fault-tolerance. Although fault-tolerance is outside the intended scope of this book, our approach in Section 6.1 blends quite well with material to be studied in Chapter 8, and in addition provides us with the opportunity to discuss a class of distributed algorithms exhibiting nontrivial stable behavior for any initial global state.

The second perspective from which we study stable properties is the perspective of stability detection, more specifically the detection of the termination of distributed computations and the detection of deadlocks. This second perspective contrasts with the first one not only because of widely differing objectives (achieving stability, in the former case, as opposed to detecting it, in the latter), but also because termination and deadlocks are far from the sort of stable properties one is seeking to achieve in the former case. Termination detection is treated in Section 6.2, where we discuss techniques for detecting the termination of distributed computations in general and of distributed computations that are of the diffusing type. These, as we will see, are characterized by the fact that $N$ is a singleton.

In Section 6.3, we discuss the detection of deadlocks in a distributed computation. Because deadlocks can occur in a variety of situations, and under assumptions that differ widely from one case to another, in Section 6.3 we concentrate on a distributed computation that controls the providing of services by the nodes to one another. Such a computation, as we describe it in that section, is deadlock-prone. The algorithm that we provide to detect the occurrence of deadlocks is very elegantly contrived, and moreover allows techniques that we have seen previously in the book, chiefly in Chapter 4, to be exercised. Exercises and bibliographic notes appear, respectively, in Sections 6.4 and 6.5.

Before we proceed to the remaining three major sections of this chapter, it may be instructive to once again return to the issue of implicit knowledge introduced in Section 2.3 for another example. Quite simply, if $P$ is a sentence related to some stable property in some global state, then $P$ is implicit knowledge that $N$ has in that global state and in all global states in its future, that is, $I_N P$. So, for example, an algorithm that has terminated or deadlocked is such that $N$ has implicit knowledge of either condition. In these cases, what the detection procedures studied in Sections 6.2 and 6.3 do is to turn such implicit knowledge into knowledge by one or more individual nodes.

### 6.1 Self-stabilization

If a distributed algorithm over $G$ can be guaranteed to lead $G$ to a global state where a particular stable property holds regardless of the global state at which the computation starts out, then the system comprising $G$ and this distributed algorithm is said to be a *self-stabilizing system*. Every self-stabilizing system is fault-tolerant in the following sense. If the local states of nodes are allowed to change infrequently as the result of a failure, then by definition the system recovers from that failure by reaching a global state at which the desired stable property is once again valid. Just how infrequent such failures have to be for self-stabilization to be still guaranteed is of course an issue, but for our purposes it suffices to recognize that failures have to be infrequent enough for the system to reach stability again once it has been disturbed.

Self-stabilizing systems do not need to be initialized, because by definition the stable property that the distributed algorithm seeks to achieve is certain to be reached from any initial global state. Also, because of the fault-tolerance connotation that inevitably accompanies the subject, once started at some initial global state, the distributed computation is supposed to be infinite, in the sense of never terminating. For this reason, not every stable property is meaningful in the context of self-stabilization, as one is interested in computations that do useful
work despite the initial state and occasionally corrupted local states. Stable properties such as global termination and deadlocks are then naturally ruled out. In order to close in on the subject more objectively, we consider the following example. Suppose that the nodes in $G$ need to utilize certain resources for their computations, but such resources cannot be utilized concurrently by any two nodes. In the context of self-stabilization, the task is to devise a distributed algorithm that, starting at any global state and given the possibility that local states may be occasionally corrupted, guarantees that the system eventually reaches a global state in which (and in whose future) no two nodes access the shared resources concurrently. (The reader may wish to check Section 8.1 for a more thorough treatment of this problem, although in that section self-stabilization is not an issue.) The stable property at hand is then that no two nodes access shared resources concurrently, so long as this can be guaranteed to remain true once it becomes true.

Henceforth in this section, $G$ is an undirected ring with FIFO edges. Referring back to the terminology of Section 2.2, the edges incident to node $n_i$ are called left and right, and the ring is assumed to be locally oriented (employing these edge denominations is only for notational convenience, though, because in this section the issue of anonymity is unimportant). Associated with a node $n_i$ is a variable $v_i$. The right of a node to access the shared resources depends on the value of its variable and on the values of the variables of its neighbors. The task of a self-stabilizing computation on $G$ is then to assign values to all nodes' variables so that no two nodes have such a right concurrently from a certain global state onward, regardless of the initial global state (i.e., the initial assignment of values to the variables).

Although a justification of this fact falls outside the scope we have intended for this book, for a ring of arbitrary size no self-stabilizing solution exists employing the exact same algorithm for all nodes. For this reason, in the solution that we present next the behavior of $n_i$ is distinct from that of the other nodes. Our solution is given as Algorithm A-Self-Stabilize, and is essentially the following. Every node $n_i$ initiates by sending the value of $v_i$ on right. Upon receiving a value $v$ on edge left, $n_i$ checks whether $v_i \neq v$. In the affirmative case, $n_i$ accesses the shared resources, and after using them sets $v_i$ to $v$ and sends $v_i$'s new value on right. An exception to this behavior is the case of $n_1$, which accesses the shared resources if $v_1 = v$ and then sets $v_1$ to $v + 1$ before sending the new value on right.

At node $n_i$, a Boolean variable $initiated_i$, initially set to false, is used to indicate whether $n_i \in N_0$. The simple behavior we just described is all there is to the algorithm, except for the possibility of faults that may corrupt a node's local state (i.e., the value of its variable). Before explaining how the algorithm is augmented to handle such faults, a few assumptions on their nature are in order. First we assume that actions do indeed take no time to be performed (cf. Section 3.2.2), and consequently the occasional faults that may corrupt a node's local state can only occur in the intervals between successive actions at that node. Faults occurring prior to the first action at a node are immaterial (because the algorithm is intended to be insensitive to initial conditions), and then it is reasonable to assume that faults can only occur at $n_i$ if $initiated_i = true$ (as in previous occasions when dealing with asynchronous algorithms, $initiated_i$ is set to true immediately upon arrival of the first message if $n_i \notin N_0$). Finally, another assumption is that $n_i$ is capable of detecting the occurrence of a fault that may have corrupted its local state. This detection is modeled as the arrival of a fault message (of purposefully undetermined origin).

The handling by $n_i$ of a fault that may have corrupted its local state goes as follows. In response to the fault message, $n_i$ sets $initiated_i$ to true (since by assumption it must have been true prior to the occurrence of the fault) and then reproduces the flow of messages involving $n_i$ during initiation. That is, it sends the value of $v_i$ (which may or may not be the same as before the occurrence of the fault) on right, and a
query message on left, the latter meant to prompt its corresponding neighbor to send the value of its variable to \( n_i \).

### Algorithm A: Self Stabilize:

**Variables:**

- \( \text{initiated}_i = \text{false} \);
- \( v_i \);

**Listing 6.1**

**Input:**

\( \text{msg}_i = \text{nil} \).

**Action if** \( n_i \in N_0 \):

- \( \text{initiated}_i := \text{true} \);
- Send \( v_i \) on right.

**Listing 6.2**

**Input:**

\( \text{msg}_i = \text{fault} \).

**Action:**

- \( \text{initiated}_i := \text{true} \);
- Send \( v_i \) on right;
- Send \( \text{query} \) on left.

**Listing 6.3**

**Input:**

\( \text{msg}_i = \text{query} \).

**Action:**

- Send \( v_i \) on right.

**Listing 6.4**

**Input:**

\( \text{msg}_i = v \).

**Action if** \( n_i \in n_1 \):

- If \( v_i = v \) then
  - begin
    - Access shared resources;
    - \( v_i := v_i + 1 \);
    - Send \( v_i \) on right;
  - end.
The first fact to notice with respect to Algorithm $A_{Self\_Stabilize}$ is that we have not assigned any initial value to $v_i$ for $v_i \in N$, precisely because of the intended insensitivity to the initial global state. Secondly, it should be noticed that (6.1) and the pair consisting of (6.2) and (6.3) are meant to be executed upon initiation, triggered respectively by the spontaneous initiation by $n_i$ if it is in $N_0$ and by the detection by $n_i$ of a the occurrence of a fault that may have corrupted the value of $v_i$. As we remarked previously, (6.2) and (6.3) are supported by our assumptions on the nature of such faults, in the sense that the response to a fault may be thought of as a re-initiation of the algorithm as far as $n_i$ is concerned.

The necessary asymmetry that we alluded to earlier is reflected in Algorithm $A_{Self\_Stabilize}$ in (6.4) and (6.5), representing respectively the action that $n_1$ and $n_i \in \{n_2, \ldots, n_n\}$ take upon receipt of a variable's value on the ring. It follows easily from these two actions that, if initially all variables have the same value, then by (6.4) $v_1$ is incremented and by (6.5) its new value is propagated on the FIFO edges around the ring until all variables have this same value. Then $v_1$ is incremented again, and so on. If, on the other hand, at least two variables have distinct values in any global state, then either the value of $v_1$ or that of $v_1 + 1$ (if $v_1 = v_n$ in that global state) is propagated on the ring as well, until $v_n$ becomes equal to $v_1$, and then the process continues repeatedly. So, although in the latter case the shared resources may be concurrently accessed by more than one node during a transient phase of some global states, a global state in which (6.4) and (6.5) cannot be executed concurrently by any two nodes is certain to occur, the same property holding for all global states in its future.

The solution by Algorithm $A_{Self\_Stabilize}$ can be turned into a solution by finite-state nodes by doing additions modulo $V$ in (6.4), so that variables are confined to the range $\{0, \ldots, V - 1\}$. Any $V$ strictly larger than $n$ will do, so that the range of values for a variable contains at least the set $\{0, \ldots, n\}$ (cf. Exercise 1).

### 6.2 Termination detection

The issue of algorithm termination appeared in this book as early as in Chapter 1, where, in Section 1.4, Algorithm Task_t runs until "global termination is known to t." As we discussed in that section, what is meant by this is that task $t$ must execute its disjunction of guarded commands until it is signaled, by means of messages that it receives, that no further messages will ever reach it and it may therefore cease executing the guarded commands and terminate its computation. The notation used in Algorithm Task_t was later modified to emphasize the reactive character of the algorithm, so that in the resulting template algorithms (Algorithms $A_{Template}$ and $S_{Template}$) only the atomic actions corresponding
to a task’s response to the receipt of messages appear. Such messages, of course, should include those intended to convey to the information that it may terminate.

Tasks have since been called nodes, and in none of the algorithms we have seen so far (or will see in chapters still ahead in the book) have we included actions to handle the treatment of the termination-related messages we have from Chapter 1 learned to be important. There are essentially two reasons why we have delayed such a treatment until this far into the book. The first reason is that global termination, as we will shortly see, is clearly an instance of stable properties, so that placing its treatment elsewhere in the book might seem a little unnatural. Secondly, and more importantly, the techniques we investigate in this section build naturally on top of what we saw in Chapters 4 and 5, often explicitly, but also sometimes simply in terms of the maturity of reasoning one must have acquired by studying those chapters.

Of course, for some of the algorithms we have seen, the issue of termination is a trivial one. For example, all the synchronous algorithms we have investigated terminate when a certain number of pulses have gone by. Similarly, in the case of all the asynchronous algorithms we have seen so far, a node should have no problem detecting that messages need no longer be expected, mostly because those algorithms are all very well structured and have very great regularity. For example, it is clear that Algorithm $A_{PI}$ terminates at a node when that node has received $inf$ from all of its neighbors, at which time it can be certain that no further message related to that algorithm will ever reach it again. Similar lines of reasoning apply to all the other asynchronous algorithms we have seen (cf. Exercise 2), as well as to many of the algorithms yet to be seen in the book. For asynchronous algorithms lacking the regularity that allows such simple termination analyses, however, the issue of detecting global termination with the purpose of relieving the various nodes from having to be on the lookout for new messages needs to be addressed from a general perspective. Asynchronous algorithms like these appear, for example, in Section 7.2.3.

The remainder of Section 6.2 is dedicated exclusively to asynchronous algorithms, although for various synchronous computations (e.g., those in Section 7.2.2) the detection of termination is not as straightforward as it has been with some of the other synchronous algorithms we have seen so far. However, the central issue in treating the termination of such algorithms is that, if they do indeed terminate, then it is essentially possible to detect that by counting pulses. Clearly, such a statement has no clear counterpart in the asynchronous case, thence our emphasis henceforth.

What we do in the next two sections is essentially to provide the atomic actions to make up for the treatment of global termination in asynchronous algorithms that do not exhibit enough regularity for its termination to be treated without messages related explicitly to termination. These actions complement those of the asynchronous algorithms proper so that the resulting asynchronous algorithms behave as intended and in addition are also capable of terminating properly. It should be clear to the reader that the techniques we describe henceforth are also applicable to synchronous algorithms exhibiting high regularity, although of course in such cases they are totally superfluous and the resulting algorithm can in all likelihood be simplified back to the one whose regularity is enough to indicate termination. Section 6.2.1 is dedicated to the case of general computations, in the sense that $N_0$ may be any subset of $N$. Section 6.2.2, on the other hand, is specific to the case in which $N_0$ is a singleton. Before entering specifics in either section, however, we must formalize a little further our concept of global termination.

An asynchronous algorithm is said to have terminated globally or reached global termination at a certain global state if every node is idle and all edges are empty in that global state. A node is idle when it is not executing any of the actions that specify its participation in the algorithm. Obviously, then, global termination is indeed a stable property, owing essentially to the reactive character of all the asynchronous computations we treat in this book. What a node needs to detect in order to be able to terminate its computation at a given local state is that, in every possible global state in which it participates with that local state, the edges on which it receives messages are all empty. Such a detection may be achieved in a variety of ways. In the case of Algorithm $A_{PI}$, for example, as soon as $inf$ has been received from all of a node’s neighbors, that node enters a local state with which it can only participate in global states that have empty edges leading to itself, and then it may terminate. When this conclusion cannot be reached in such a straightforward manner, additional computation
The distributed computation of interest in this section is initiated by any subset \( N_0 \) of \( N \) and progresses through the exchange of messages generically referred to as \( \text{comp}-\text{msg} \)’s. We take \( G \) to be a strongly connected directed graph, so that the case of an undirected \( G \) can also be handled in a straightforward manner. Our approach to termination detection in this section is based strongly on Algorithm \( \text{A\_Record\_Global\_State} \), and then we assume that \( G \)'s edges are FIFO.

The approach we take goes essentially as follows. Before going idle, a node that "suspects" it may have terminated initiates the recording of a global state. This suspicion is of course highly dependent upon the particular computation at hand, so we let it be indicated by a Boolean variable \( \text{suspects} \) at \( n_i \in N \). This variable is set to either \( \text{false} \) or \( \text{true} \) after, in accordance with Algorithm \( \text{A\_Template} \), \( n \) has computed and possibly sent out some messages, either spontaneously if \( n, N_0 \) or upon the receipt of a \( \text{comp}-\text{msg} \) (the initial value assigned to the variable is then unimportant). A global state in which \( \text{suspects}=\text{true} \) for all \( n_i \in N \) does not imply global termination in that global state (because there may be messages in transit), but we assume that global termination does imply that \( \text{suspects}=\text{true} \) for all \( n_i \in N \).

This recording of a global state proceeds entirely along the lines of Algorithm \( \text{A\_Record\_Global\_State} \), that is, through the exchange of marker messages, and may as in that case be initiated concurrently by more than one node if for such nodes the \( \text{suspects} \) variables become \( \text{true} \) concurrently. When the recording of a global state is completed at \( n_i \), it then sends what it recorded to \( n_i \) (the assumed leader), which, upon receiving similar information from every node, checks whether the global state that was recorded indicates global termination. If it does, then a terminate message is broadcast by \( n_i \) to all of \( G \)'s nodes, which then terminate.

Clearly, this procedure may be wasteful because a node \( n_i \) that receives a marker when \( \text{suspects}=\text{false} \) should not propagate the marker’s onward because the resulting global state cannot possibly indicate global termination. Aborting a global state recording is not something we have considered before, and there are a few implications to be considered. In our present context, the two problems that result from prematurely aborting a global state recording are the need to terminate the aborted recording properly and the possibility that \( n_i \) receives incomplete global states which must somehow be dealt with. We tackle both problems simultaneously, as follows. Every marker is sent with a tag, and every node keeps record of the greatest tag it has seen so far in a marker. If the tag a node attaches to a marker is sent out when initiating a new global state recording is strictly greater than any tag it has ever seen, then the rule for participating in global state recordings is very simple. A node only participates in a new global state recording if the tag accompanying the corresponding marker is strictly larger than every tag it has known of and in addition its \( \text{suspects} \) variable is \( \text{true} \). Any marker received in different circumstances is ignored. The reader should note that this provides the necessary control for terminating aborted global state recordings, and also allows \( n_i \) to discard useless information it has recorded if the information that it receives from nodes on a recorded global state is itself accompanied by the tag that was attached to the marker’s during the recording of that global state.

This strategy is adopted by Algorithm \( \text{A\_Detect\_Termination} \), given next. In addition to the Boolean variable \( \text{suspects} \), node \( n \) needs some of the variables employed by Algorithm \( \text{A\_Record\_Global\_State} \) as well. These are initialized to \( \emptyset \), for each node.
node $n_i \in I_{\text{Neig}}$, and the Booleans $\text{recorded}_i$, and $\text{received}_i$, all initialized to $\text{false}$. The maximum tag $n_i$ has seen in a marker is denoted by $\text{max-tag}_i$. Finally, another Boolean variable, $\text{terminated}_i$, initially equal to $\text{false}$, is used by $n_i$ to indicate whether a terminate message has been received from $n_i$ ($\text{terminated}_i$ is used to indicate that global termination has been detected). This variable, in Algorithm Task_\text{t}, can be used to exit the repeat...until loop.

In Algorithm A\_Detect\_Termination, marker messages are sent as marker($t$) messages, where $t$ is a positive tag. The initial value of $\text{max-tag}_i$ is then zero.

**Algorithm A\_Detect\_Termination:**

**Variables:**

- $\text{suspects}_i$
- $\text{edge\_state}_i = \emptyset$ for all $n_i \in I_{\text{Neig}}$
- $\text{recorded}_i = \text{false}$
- $\text{received}_i = \text{false}$ for all $n_i \in I_{\text{Neig}}$
- $\text{max\_tag}_i = 0$
- $\text{terminated}_i = \text{false}$

Other variables used by $n_i$ and their initial values, are listed here.

**Listing 6.6**

**Input:**

$\text{msg}_i = \text{nil}$.

**Action if** $n_i \in N_0$:

- Do some computation;
- Send one $\text{comp\_msg}$ on each edge of a (possibly empty) subset of $\text{Out}_i$;
- if $\text{suspects}_i$ then
  - begin
    - $\text{max\_tag}_i := \text{max\_tag}_i + 1$
    - $\text{recorded}_i := \text{true}$
    - Send marker($\text{max-tag}_i$) to all $n_j \in O_{\text{Neig}}$
  - end.

**Listing 6.7**

**Input:**

$\text{msg}_i = \text{comp\_msg}$ such that $\text{origin}_i(\text{msg}_i) = (n_j \rightarrow n_i)$.

**Action:**

- Do some computation;
- Send one $\text{comp\_msg}$ on each edge of a (possibly empty) subset of $\text{Out}_i$;
- if $\text{recorded}_i$ then
  - if not $\text{received}_i$ then
if suspects, then
begin

\[ \text{edge state}^j_i \quad \text{edge state}^k_i \cup \{\text{msg}\}; \]

begin

\[ \text{edge state}^k_i \; : = \; \emptyset \text{ for all } n_k \in I_{\text{Neig}}; \]

\[ \text{received}^k_i \; : = \text{false} \text{ for all } n_k \in I_{\text{Neig}}; \]

\[ \text{max tag} \; : = \text{max tag} + 1; \]

\[ \text{recorded} \; : = \text{true}; \]

Send marker(max_tag) to all \( n_i \in O_{\text{Neig}}. \)

end.

\[ \text{Listing 6.8} \]

**Input:**

\( \text{msg} = \text{marker}(t) \) such that \( \text{origin}({\text{msg}}) = (n_i \rightarrow n_j). \)

**Action:**

if \( t = \text{max tag} \), then

\[ \text{received}^j_i \; : = \text{true}; \]

if \( t > \text{max tag} \), then

begin

\[ \text{max tag} \; : = t; \]

\[ \text{edge state}^j_i \; : = \emptyset \text{ for all } n_k \in I_{\text{Neig}}; \]

\[ \text{recorded} \; : = \text{false}; \]

\[ \text{received}^k_i \; : = \text{false} \text{ for all } n_k \in I_{\text{Neig}}; \]

if suspects, then

begin

\[ \text{received}^j_i \; : = \text{true}; \]

\[ \text{received}^k_i \; : = \text{true}; \]

Send marker(max_tag) to all \( n_k \in O_{\text{Neig}}. \)

end

end;

if

\[ \text{received}^k_i \text{ for all } n_k \in I_{\text{Neig}}, \text{then} \]

Send \( \text{edge state}^k_i \) for all \( n_k \in I_{\text{Neig}}, \) along with max_tag, to \( n_i. \)
This algorithm is, in essence, a blend of Algorithm \texttt{A_Template} on \texttt{comp_msg}’s and Algorithm \texttt{A_Record_Global_State}. Specifically, (6.6) is \((2.1)\) enlarged by \((5.11)\) to initiate a global state recording if \texttt{suspects} \(_i\) = \texttt{true}. Similarly, \((6.7)\) is \((2.2)\) enlarged by \((5.11)\) and \((5.13)\), respectively to initiate a global state recording if \texttt{suspects} \(_i\) = \texttt{true} and to record the messages that comprise an edge’s state in the global state being recorded. Finally, \((6.8)\) is \((5.12)\), conveniently adapted to abort ongoing global state recordings upon receipt of a \texttt{marker} carrying a tag strictly greater than the greatest one the node has seen.

Let us consider the functioning of Algorithm \texttt{A_Detect_Termination} more carefully. Node \(n_i\) performs computation, possibly with the sending of some \texttt{comp_msg}’s, either spontaneously, if it is a member of \(N_0\), or upon receiving a \texttt{comp_msg}. In the former case, \(n_i\) may also start a global state recording with \texttt{marker}’s carrying a tag increased by one with respect to the greatest tag it has seen (cf. (6.6)). In the latter case, \(n_i\) may record the \texttt{comp_msg} as part of the state of the edge on which it was received, or it may, as in the other case, start the recording of a global state, after re-initializing its variables related to the recording of global states, or it may do both, in which case the recording of \texttt{comp_msg} will have been in vain (cf. (6.7)). Notice that these two possibilities account for all the opportunities \(n_i\) has to start a global state recording, and in both cases such a start is done with properly initialized variables. Because \(n_i\) maintains only one set of variables for global state recording, it may only participate in the recording of one global state at a time, so that upon initiating its participation in a new recording it must quit its participation in whatever recording it may have been participating so far. This is the reason for variable re-initialization in \((6.7)\) if \texttt{suspects} \(_i\) becomes \texttt{true}.

The other occasion in which \(n_i\) may have to forsake its current participation in a global state recording is upon receiving a \texttt{marker} \((t)\) such that \(t > \texttt{max_tag}\). When this happens, \(n_i\) re-initializes its variables related to global state recording and, if \texttt{suspects} \(_i\) = \texttt{true}, joins in the new global state recording, as in \((6.8)\). If \(t < \texttt{max_tag}\), then the \texttt{marker} \((t)\) is ignored (it clearly belongs to a long-forsaken global state recording), whereas if \(t = \texttt{max_tag}\), then it may correspond to a recording in which \(n_i\) is currently engaged. The receipt of a marker in \((6.8)\) may also imply that \(n_i\) has finished its participation in the current global state recording, and then what it recorded is sent to the leader for analysis (it only sends the edge states, though, because by \((6.6)\) through \((6.8)\) \(n_i\) does not participate in the recording of a global state if its local state is anything other than \texttt{suspects} \(_i\) = \texttt{true}). This information is sent to the leader along with the tag with which it was recorded, and the leader, upon having received information with the same tag from all nodes, decides whether global termination has been reached, in which case a \texttt{terminate} message is broadcast to all nodes. We have omitted from the algorithm the actions for \(n_i\) to perform its role as a leader, and we have also omitted any specific mention to how the broadcast of the \texttt{terminate} order is performed. Filling in these blanks should pose no difficulty, though, especially after our discussion of information propagation in \texttt{Section 4.1} (cf. Exercise 3). The response of \(n_i \neq n_1\) to the \texttt{terminate} message is in \((6.9)\).

The correctness of Algorithm \texttt{A_Detect_Termination} is based on \texttt{Theorem 5.5} on the correctness of Algorithm \texttt{A_Record_Global_State} if \(G\) is strongly connected with FIFO edges, and on the following observation. Suppose a global state in which global termination holds does exist. As we assumed earlier in this section, at this global state it must hold that \texttt{suspects} \(_i\) = \texttt{true} for all \(n_i \in N\), so we may consider the greatest value of \texttt{max_tag} over all of

\begin{verbatim}
Listing 6.9

Input:
\texttt{msg} = \texttt{terminate.}

Action if \(n_i \neq n_1;\)
\texttt{terminated} := \texttt{true}.
\end{verbatim}

When the corresponding suspects’s became true for the last time. The nodes at which this greatest value occurred must by (6.6) and (6.7) have initiated a global state recording concurrently, and by (6.8) this global state recording must have been propagated by all nodes. Consequently, at least one global state recording is carried out to completion, including the recording of a global state in which global termination holds.

Before leaving this section, a couple of observations are worth making. The first observation concerns obvious possible simplifications to Algorithm A_Detect_Termination, especially in what concerns the reports that are sent to \( n_i \). We elaborate no further on the issue, but encourage the reader to further investigate it (cf. Exercise 4). The second observation relates to the treatment of computations for which global termination does not hold at any global state. Computations like these appear in Section 10.2, and because they ordinarily do not terminate by themselves, what we seek is to force their termination by detecting termination-related properties that appear in some global states with special characteristics. As we discussed briefly in Section 5.2.2, in this case a leader can be employed to search for the special global states, and upon finding one of them for which the desired termination-related properties do hold the leader then directs all other nodes to terminate. In Chapter 10, we address these issues with more detail.

### 6.2.2 Diffusing computations

In this section, we concentrate on detecting the termination of asynchronous algorithms for which \( N_0 \) has one single member, assumed to be \( n_i \), the leader. \( G \) is in this section taken to be an undirected graph. The approach we described in the previous section is of course applicable to this case as well (and, for that matter, so is the approach of this section applicable to cases in which \( N_0 \) is not a singleton—cf. Exercise 5), although in that case \( n_i \) would no longer be required to be a member of \( N_0 \).

Distributed computations for which \( N_0 \) is a singleton are referred to as diffusing computations, because in such computations the causality that the flow of messages induces is "diffused" from one single node. Of course this same intuition is also present in the cases of larger sets of initiators, but the denomination as a diffusing computation is not generally used in those cases because it would seem unnatural to say that the computation is diffused from the members of \( N_0 \) when such a set can be arbitrarily large, possibly equal to \( N \).

The algorithm we saw in Section 4.1.1 to propagate information with feedback from \( n_i \) is an example of diffusing computations, and, as we will shortly see, it is an example of particular interest in the context of detecting the termination of diffusing computations in general. In Algorithm A_PIF, the role played by \( n_i \) can be thought of as being not only that of the original propagator of int, but also that of the detector of when the propagation has terminated throughout all of \( G \). Although, as we remarked earlier in Section 6.2, in that algorithm every node can decide upon its termination rather easily (without the need for intervention from \( n_i \)), the general idea of having a wave of information collapse back to \( n_i \) upon global termination is quite useful for computations whose termination cannot be detected so simply.

One of the main motivations to look for a different solution in the case of diffusing computations, rather than just employ the general technique of the previous section, is the potentially very high complexity of the methodology realized by Algorithm A_Detect_Termination. Although, due to its generality, we did not attempt any analysis when presenting that algorithm, clearly its complexity depends on the number of global state recordings it performs, so that the overall complexities may be too high. The specialized solution we study in this section, on the other hand, allows global termination to be detected without affecting the complexities of the computation proper.

The following is an outline of Algorithm A_Detect_Termination_D ("D" for Diffusing) . Every comp_msg is acknowledged with an ack message. Node \( n_i \) maintains a counter expected, initially equal to zero, to indicate the number of ack messages \( n_i \) receives from its neighbors (we assume that expected is automatically increased whenever a comp_msg is sent by \( n_i \)). As in the case of Algorithm A_PIF, \( n_i \) also maintains a variable parent, initialized to nil, to indicate the origin of a comp_msg received in a special situation to be described shortly. The behavior of \( n_i \) is then the following. Whenever \( n_i \) receives a comp_msg and expected > 0, an ack is immediately sent in response. If, on the other hand, a comp_msg is received and expected = 0, then the ack is withheld and sent only when expected becomes equal to zero again (if it at all changes with the computation \( n_i \) does in response to the arriving comp_msg,
otherwise the *ack* is sent immediately after that computation). The variable *parent* is in this case set to point to the node that sent the *comp_msg* until the *ack* can be sent. We say that *n* has reached a state of *tentative termination*, or that *n* has *tentatively terminated* when *expected* becomes zero and the pending *ack*, if any, is sent to *parent*. This condition may, however, change many times during the computation, for *expected* may again acquire a positive value as a consequence of the reception of a *comp_msg*. Global termination is detected when *n* has tentatively terminated, which in the case of *n* may happen only once.

The resulting algorithm is Algorithm A._Detect_Termination_D_, presented next. As in the previous section, a *terminate* message is employed by *n* to broadcast the detection of global termination. A Boolean variable *terminated*, initially set to *false*, is employed by *n* to signal that *n* may exit the *repeat* … *until* loop in Algorithm Task_t. This variable is set to *true* by *n* upon detection of global termination, if *n* = *n*1, or upon receipt of the *terminate* message, otherwise.

**Algorithm A._Detect_Termination_D_:**

- **Variables:**
  - *expected* = 0;
  - *parent* = nil;
  - *terminated* = *false*.

**Listing 6.10**

**Input:**
- *msg* = nil.

**Action** if *n* ∈ *N*0:
- Do some computation;
- Send one *comp_msg* on each edge of a (possibly empty) subset of *Inc*.

**Listing 6.11**

**Input:**
- *msg* = *comp_msg* such that *origin*(*msg*) = (*n*, *nj*).

**Action:**
- if *expected* > 0 then
  - begin
    - Send *ack* to *nj*;
    - Do some computation;
    - Send one *comp_msg* on each edge of a (possibly empty) subset of *Inc*;
  - end
- else
  - begin
    - Do some computation;
    - Send one *comp_msg* on each edge of a (possibly empty) subset of *Inc*;
    - if *expected* > 0 then
      - *parent* := *nj*
    - else
      - Send *ack* to *nj*
In Algorithm $A_{\text{Detect\_Termination\_D}}$, (6.10) and (6.11) are, in essence, (2.1) and (2.2), respectively, in Algorithm $A_{\text{Template}}$ on comp_msg's, while (6.12) and (6.13) deal with the reception of ack and terminate messages, respectively (the latter for $n \neq n_1$). Together, (6.11) and (6.12) can be seen to be closely related to (4.4) in Algorithm $A_{\text{PIF}}$ in that all of them are involved with withholding acknowledgements from a parent neighbor until it is appropriate for that acknowledgement to be sent. This similarity with those two algorithms allows Algorithm $A_{\text{Detect\_Termination\_D}}$ to be interpreted as a general template for asynchronous diffusing computations in which $n_1$, the computation's sole initiator, detects global termination upon being reached by a collapsing wave of acknowledgements. This view of a computation as a propagating wave is the same that we employed in various occasions in Chapter 4, and in the present context allows the following pictorial interpretation. In Algorithm $A_{\text{Detect\_Termination\_D}}$, a wave is initiated by $n_1$ in (6.10) and throughout $G$ it propagates back and forth with respect to $n_1$. It propagates away from $n_1$ with comp_msg's and backwards in the direction of $n_1$ with ack's. When the wave hits $n_i$ in its forward propagation, it may bounce back immediately (if expected; $> 0$ at the beginning of (6.11) or expected; $= 0$ at the end of (6.11)) or it may continue further on from that node (otherwise). Node $n_i$ may in this case be $n_1$ itself, in which case the wave is sure to bounce back at once. The wave that propagates backwards in the direction of $n_1$ does so by means of ack messages, and continues to propagate at each node $n_i$ that it encounters so long as expected; becomes zero with its arrival. What differentiates the wave propagations in this case from those of Algorithm $A_{\text{PIF}}$ is that a node that has already seen the ack wave go by may be hit by a forward-moving wave again (that is, by a comp_msg), so that overall the picture is that of a wave that may oscillate back and forth several times, and in different patterns on the various portions of $G$, before it finally collapses back onto $n_1$.

Before proceeding with a more formal analysis of this behavior, we mention that, as in the case of Algorithm $A_{\text{Detect\_Termination\_D}}$ of the previous section, we have not in Algorithm $A_{\text{Detect\_Termination\_D}}$ been complete to the point of specifying the termination of $n_1$ and
the propagation of the \textit{terminate} broadcast. The reader should work on providing the missing
details (cf. Exercise 6).

The correctness of Algorithm \texttt{A\_Detect\_Termination\_D} is established by the following
theorem.

\textbf{Theorem 6.1.}
\textit{Every global state in which }$n_1$\textit{ has tentatively terminated in Algorithm
\texttt{A\_Detect\_Termination\_D} is a global state in which global termination holds.}

\textbf{Proof} If $n_1$ has tentatively terminated, then by (6.11) and (6.12) every node must have sent a
finite number of \texttt{comp\_msg}'s. As these \texttt{comp\_msg}'s and the corresponding \texttt{ack}'s were
received, the value of \texttt{expected} for node $n_i$, initially equal to zero, became positive and zero
again, possibly several times. Whenever a transition occurred in the value of \texttt{expected} from
zero to a positive value, \texttt{parent} was set to point to the node that sent the corresponding
\texttt{comp\_msg}. Consider the system states in which every node $n_i$ is either in a state of positive
\texttt{expected}, following the last transition from zero of its value, if it ever sent a \texttt{comp\_msg} during
the diffusing computation, or in any state, otherwise. Clearly, at least one of these system
states is a global state, as for example the one in which every node that ever sent
\texttt{comp\_msg}'s is in its state that immediately precedes the reception of the last \texttt{ack} (Figure
6.1). In this global state, only \texttt{ack}'s flow on the edges, none of which sent as a consequence
of the reception of a last \texttt{ack}. Let us consider one of these global states.

In this global state, the variables \texttt{parent} for $n_i \neq n_1$ induce a tree that spans all nodes in $G$
corresponding to nodes that sent at least one \texttt{comp\_msg} during the diffusing computation.
(This tree is in fact dynamically changing with the progress of the algorithm, as \texttt{parent} may
point to several of $n_i$'s neighbors along the way; it is always a tree, nevertheless.) This tree is
rooted at $n_1$, and its leaves correspond to those nodes from which no other node $n_i$ received
the \texttt{comp\_msg} that triggered the last transition from zero to a positive value of \texttt{expected}. As
in the proof of \textbf{Theorem 4.1}, we proceed by induction on the subtrees of this tree. Along the
induction, the assertion to be shown is that every global state in which the subtree's root has
tentatively terminated is a global state in which every other node in the subtree has also
tentatively terminated.

The basis of the induction is given by the subtrees rooted at the leaves, and then the
assertion clearly holds, as no leaf $n_i$ is such that $n_i = \texttt{parent}$, for some
Figure 6.1: Edges in the precedence graph fragment shown in part (a) are drawn as either solid lines or dashed lines. Solid lines represent comp\_msg's, dashed lines represent ack's, and the remaining edges of the precedence graph are omitted. In this case, system\_state\(>(\Xi_1, \Xi_2)\) is clearly a global state, and is such that every node that ever sent a comp\_msg during the diffusing computation (i.e., \(n_1\) and \(n_3\)) is in the state that immediately precedes the reception of the last ack. In part (b), the spanning tree formed by the variables parent, for each node \(n_i\) in this global state is shown with directed edges that point from \(n_i\) to \(n_j\) to indicate that parent\(=n_j\). In this case, the tree has \(n_1\) for root and its single leaf is \(n_3\).

As the induction hypothesis, assume the assertion for all the subtrees rooted at nodes \(n_j\) such that parent\(=n_j\). Then \(n_i\) receives expected\_ack’s, at which time it has tentatively terminated, and by the induction hypothesis so have all other nodes.

Let us now return briefly to the question, raised earlier in this section, of the algorithm’s complexities. Because exactly one ack is sent for each comp\_msg, the message complexity of Algorithm A\_Detect\_Termination\_D is exactly the message complexity that Algorithm A\_Template would have to realize the same computation without having to detect global termination. The same holds with respect to the algorithms’ time complexities, because the time that Algorithm A\_Detect\_Termination\_D spends in addition to that already spent by the corresponding instance of Algorithm A\_Template is used solely for the final collapsing of the ack wave onto \(n_1\). This additional time, clearly, does not exceed that of Algorithm A\_Template, as this wave that propagates backwards comes from as far as the corresponding forward-propagating wave got.

6.3 Deadlock detection
Deadlocks are a very close acquaintance of anyone who has been involved with any of the many facets of concurrency at any depth. In this book, our concern for deadlock situations has already appeared explicitly in a couple of places, as in Sections 1.3 and 1.5, and less conspicuously it has also appeared in some other situations, as for example in our discussion on the importance of distinct identifications for nodes in the context of leader
Informally, a group of nodes is in *deadlock* when every node in the group is suspended for a condition that can only be realized by nodes that belong to the group as well. Clearly, then, deadlocks are indeed stable properties. The classical approaches to the treatment of deadlocks range from its prevention (as for example in Sections 1.3 and 1.5) to its detection after it has occurred. The prevention of deadlocks is based on making sure, by design, that at least one of the conditions necessary for the occurrence of deadlocks can never hold. One of these conditions is the so-called *wait cycle*, which in our context consists of a subset of \( N \) whose members are cyclically waiting for one another. Forbidding the occurrence of such cycles constitutes the strategy we described in Section 1.5 to prevent deadlocks related to message buffering (cf. Theorem 1.1).

The detection of deadlocks, on the other hand, is based on the rationale that it may be simpler, or less restrictive in a variety of senses, not to impose conditions leading to the prevention of deadlocks, but rather to let them occur occasionally and then proceed to detecting them when the suspicion exists that they may have indeed occurred. Because deadlocks are stable properties, an approach to detecting their occurrence is to record a global state of the system and then work on this global state to check for the presence of any deadlock. If a deadlock is found in the global state that was recorded, then because of its stability it must have persisted as the system continued to evolve following the recording of the global state. If, on the other hand, no deadlock was found, then naturally the only possible conclusion is that no deadlock existed in any global state in the past of the recorded global state, although it may have occurred in global states in its future.

This section is dedicated to the study of deadlock detection in the case of a very specific distributed computation. Aside from the deadlock issue *per se*, the benefits of this study are manifold. In particular, the approach we describe to deadlock detection yields a distributed algorithm to perform the detection that works on a recorded global state in a completely distributed fashion. This is in contrast with our previous use of recorded global states in this chapter, for example in Section 6.2.1, where the analysis of the recorded global state to detect the desired stable property was performed in a centralized fashion by a leader. By contrast, our approach in this section performs the detection without moving any of the recorded information from the node where it was recorded. Another benefit is that the algorithm we describe constitutes another elegant example of the wave techniques we have seen so far in the book, notably in Chapters 4, 5 and in this very chapter.

We proceed in the following two sections as follows. In Section 6.3.1, the asynchronous computation that may deadlock is introduced. In Section 6.3.2, an algorithm is given to look for deadlocks in a recorded global state of that computation. In both sections, \( G \) is taken to be an undirected graph with FIFO edges.

### 6.3.1 The computation

Every node in \( G \) is the provider of a service to some of the other nodes. An edge \((n_i, n_j)\) exists in \( G \) if and only if at least one of \( n_i \) and \( n_j \) may request the service provided by the other node. Node \( n_i \) has a Boolean variable *available*, initially set to *true*, to indicate whether it is available to provide a service it is requested or not. Because \( n_i \) may only respond to one request at a time, every request that it receives when *available* = *false* must wait to be serviced when *available*, becomes *true*.

A node requests a service to one of its neighbors by sending it a message *request*. When the request is finally honored, a message *done* is used to indicate that. In the computations that we consider, nodes are allowed to request the same service to more than one neighbor at a time. In addition, if \( n_i \) wishes to request a service to \( x_i \) of its neighbors, then it is also allowed to do the request to \( y_i \supseteq x_i \) neighbors. Upon receiving \( x_i \) done messages, it then sends a *quit* message to the \( y_i - x_i \) nodes from which it did not receive a *done*. Clearly, \( 0 < x_i \subseteq y_i \subseteq |Inc_i| \) for all \( n_i \in N \). For simplicity, we assume that \( x_i \) and \( y_i \) are constants for each \( n_i \in N \).

What accounts for the possibility of deadlocks in this computation is that a node, while servicing the request of one of its neighbors, may itself issue requests for services that it needs some of its neighbors to perform in order to finish its own task. The possibility of
deadlocks is then obvious, given that no node accepts a new request if it has a pending request itself. A node that receives a *quit* on a service for which it sent *request's* sends out *quit's* itself.

This computation is given more formally next, in the form of Algorithm *A_Provide_Service*. In this algorithm, a variable *requester*\(_i\), initially equal to *nil*, is used by \(n_i\) to point to the node to which it is currently providing service. This variable is maintained in such a way that, for \(n_i \in N_0\), \(available_i = \text{true}\) if *requester*\(_i\) = *nil* (but not conversely, so no redundancy really exists between the two variables). In addition, a set *pending*\(_i\) \(\subseteq Neig_i\) is used by \(n_i\) to keep track of the neighbors to which it sent *request* messages without however having received a *done*. This variable is such that \(y_i - x_i \subseteq | \subseteq y_i\).

**Algorithm A_Provide_Service:**

**Variables:**
- \(available_i = \text{true}\);
- \(requester_i = \text{nil}\);
- \(pending_i\).

**Listing 6.14**

**Input:**
- \(msg_i = \text{nil}\).

**Action if** \(n_i \in N_0\):
- Let \(Y_i \subseteq Neig_i\) be such that \(y_i|Y_i|\);
- Send *request* to all \(n_k \in Y_i\);
- \(pending_i := Y_i\).

**Listing 6.15**

**Input:**
- \(msg_i = \text{request}\) such that \(origin(msg_i) = (n_i, n_j)\).

**Action when** \(available_i\):
- **if** service from other nodes is needed **then**
  - **begin**
    - Let \(Y_i \subseteq Neig_i\) be such that \(y_i|Y_i|\);
    - Send *request* to all \(n_k \in Y_i\);
    - \(pending_i := Y_i\);
    - \(requester_i := n_j\);
    - \(available_i := \text{false}\)
  - **end**
- **else**
  - **begin**
    - Perform requested service;
    - Send *done* to \(n_j\)
  - **end**

**Listing 6.16**
Input:
\[ msg_i = \text{done} \] such that \( \text{origin}(msg_i) = (n, n_i) \).

Action:
\[
\begin{align*}
\text{pending}_i & := \text{pending}_i - \{n_i\}; \\
\text{if} \ |\text{pending}_i| = y_i - x_i \text{ then} & \\
\text{begin} & \\
\text{Send} \text{ quit} \text{ to all } n_j \in \text{pending}_i; \\
\text{if} \ \text{requester}_i \neq \text{nil} \text{ then} & \\
\text{begin} & \\
\text{Perform} \text{ requested service}; \\
\text{Send} \text{ done} \text{ to} \text{requester}_i, \\
\text{end}; & \\
\text{available}_i := \text{true} & \\
\text{end}. \\
\end{align*}
\]

The reader should have no difficulties to check that (6.14) through (6.17) do indeed realize the computation we outlined earlier on \( G \). Even so, it may be instructive to check the use of \( \text{available}_i \) as a condition for (6.15) to be carried out (cf. Algorithm A Template).

6.3.2 An algorithm
The possibility of deadlocks in Algorithm A_Provide_Service is very clearly visible in (6.14) and (6.15), because request messages may be sent in such a way that a wait cycle is formed in \( G \). One simple example is the situation in which some of the nodes in \( N \) send request's to one another in a cyclic fashion. Another example is the case in which a node triggers a chain of request's that ends up in itself.

When a node \( n_i \) has waited "too long" (or longer than would be "typical") for the \( x_i \) done messages that it expects to be received, it may start a deadlockdetection procedure to verify whether it is involved in a deadlock. The procedure that we describe in this section is, in much the same way as Algorithm A_PIF, designed to be started by one node only (so \( N_0 \) must be a singleton). We assume, without any loss in generality, that such a node is \( n_1 \), but it should be clear that in general all the messages related to this detection must bear an indication of which node initiated the process so that multiple concurrent detections started by different nodes do not interfere with one another.

What \( n_1 \) does to detect the occurrence of a deadlock is to start the recording of a global state, and then to start a detection procedure on the global state that was recorded. The global state is stored in the same distributed fashion as it was recorded, and the detection procedure is itself an asynchronous algorithm in which various nodes participate. However, the deadlock detection does not operate on the entirety of \( G \). Instead, this procedure runs on some portions of \( G \) given in accordance with what is known as a wait graph. The node set of
the wait graph is a subset of \( N \), and its edges are directed versions of some of the edges in \( E \). In order for a node \( n_i \) to be in the wait graph, at least one of the edges incident to it in \( G \) must also be in the wait graph.

The conditions for edges of \( G \) to be edges of the wait graph vary dynamically as the computation given by Algorithm \( A_{Provide\ Service} \) evolves. In a particular global state of that computation, an edge \((n_i,n_j)\) is an edge of the wait graph if and only if, in that global state, all of the following three conditions hold.

- \( n_i \) has sent \( n_j \) a request.
- \( n_i \) has not sent \( n_j \), a done.
- \( n_i \) has not sent \( n_j \), a quit.

These three conditions include messages that have been received as well as messages in transit. In particular, in that global state there may be a request in transit from \( n_i \) to \( n_j \), but no done in transit from \( n_j \) to \( n_i \) and no quit in transit from \( n_i \) to \( n_j \), so that what the conditions imply is that \( n_i \) has requested a service to \( n_j \), and is in that global state waiting for the service to be performed. In the wait graph, such an edge is directed from \( n_i \) to \( n_j \) to indicate precisely that wait. At \( n_i \), and in the context of a particular global state, \( out\_wait \) is the subset of \( Neig \), such that \( n_i \in out\_wait \) if and only if an edge directed from \( n_i \) to \( n_i \) exists in the wait graph in that global state. The set \( in\_wait \) is defined likewise to include those nodes \( n_i \) such that an edge directed from \( n_i \) to \( n_i \) exists.

It is on the portions of \( G \) that intersect the wait graph that the deadlock detection should run, as it is on those portions that the waiting is taking place. This should pose no problem, because the detection runs on a recorded global state and in that global state the wait graph is well defined, as we just discussed. However, the recording of the global state cannot quite run on the wait graph as well, because no such graph has yet been determined (determining it is, in fact, the very purpose of the global state recording). On the other hand, it seems clearly a waste to perform the global state recording all over \( G \), because a great portion of it may not have the slightest chance of participating in the wait graph once the global state is recorded. However, the only other appropriate structure related to \( G \) that the global state recording might utilize is that given by the sets \( pending \) for \( n \in N \) if they are nonempty (cf. Section 6.3.1), but they are not enough to describe the desired graph that would be "between" \( G \) and the wait graph to be eventually obtained (cf. Exercise 7).

So what \( n_i \) does is to initiate a global state recording over \( G \) as in Algorithm \( A_{Record\ Global\ State} \), and then to initiate a deadlock detection procedure on the wait graph, which we describe next as Algorithm \( A_{Detect\ Deadlock} \). Because all edges are FIFO, \( n_i \) might in principle initiate the deadlock detection immediately after initiating the global state recording. If this were done, then Algorithm \( A_{Detect\ Deadlock} \) would need a little extra control to ensure that a node would only participate in the latter computation after being through with its participation in the former. In order to avoid this unnecessary complication, we assume that \( n_i \) is somehow notified of the global termination of the global state recording. The reader should consider with care the design of an asynchronous algorithm to record global states and signal its initiator (assumed unique) upon the recording's global termination (cf. Exercise 8).

For each node \( n_i \), the local state to be recorded comprises the variables \( available, requester \), and \( pending \). The edge states to be recorded may in turn contain request, done, and quit messages. Once the recording is completed at \( n_i \) (i.e., \( n_i \)'s local state has been recorded and so have the states of all edges leading toward \( n_i \)), the sets \( in\_wait \) and \( out\_wait \), that describe the wait graph at \( n_i \) can be determined as follows. The set \( in\_wait \) must include the node \( requester \); (if \( available = false \) and \( n_i \) is not in the \( N_N \) of Algorithm \( A_{Provide\ Service} \)) and every neighbor \( n_j \) such that the recorded state of the edge \((n_i,n_j)\) in the direction from \( n_i \) to \( n_j \) contains a request but does not contain a quit. Similarly, the set \( out\_wait \) must include every neighbor \( n_j \) that is in \( pending \), and such that neither the recorded state of \((n_i,n_j)\) in the direction from \( n_i \) to \( n_j \) has a done nor the recorded state of \((n_j,n_i)\) in the opposite direction has a quit. It is a simple matter to check that these sets are consistent over all edges, that is, \( n_i \in out\_wait \) if and only if \( n_i \in in\_wait \), and conversely, for all edges \((n_i,n_j)\).

The following is a general outline of Algorithm \( A_{Detect\ Deadlock} \). First a wave of notify messages is propagated by \( n_i \) along the edges leading to nodes in the \( out\_wait \) sets. Because an \( in\_wait \) set may contain more than one node, this wave may reach a node more
than once and should only be sent forward upon receipt of the first *notify*. A node having an empty *out_wait* set does not propagate the *notify*'s onward, but rather starts the propagation of another wave, this time with *grant* messages and on edges leading to nodes in the *in_wait* sets. Such waves simulate the concession of the services upon which nodes wait. A node that receives as many *grant*’s as it needs (this is given by its *x* constant) propagates the wave onward on its own *in_wait* set, as in the simulation such a node has already been granted services from as many neighbors as it needs. This wave is propagated as far back as nodes with empty *in_wait* sets, from which it collapses back with *grant_done* messages. A node with an empty *in_wait* set sends a *grant_done* message immediately upon receiving a *grant*. Other nodes withhold the *grant_done* that corresponds to the *x*-th *grant*, but do respond with immediate *grant_done*’s upon receiving all other *grant*’s. At node *n*, the node from which the *grant_done* is withheld is pointed to by *out_parent*. The *grant_done* to *out_parent* is sent when *n* has received as many *grant_done*’s as there are nodes in *in_wait*. What remains now is to collapse back onto *n* the wave that it propagated with *notify* messages. This is accomplished with *notify_done* messages as follows. Node *n*, upon receiving the first *notify*, points to its sender with a variable *in_arent*. Every other notify is replied to immediately with a *notify_done* message. Whenever the nodes that initiated the *grant* waves have received as many *grant_done*’s as there are nodes in their *in_wait* sets, they send a *notify_done* to their *in_parent* neighbors. Other nodes do the same upon receiving as many *notify_done*’s as there are nodes in their *out_wait* sets. When *n* receives all the *notify_done*’s that are due, it then checks the number of *grant*’s it received along the process. Node *n* is in deadlock if and only if this number is less than *x*. In this algorithm, nodes behave as if they could grant service concomitantly for all the requests they receive. This is of course untrue by assumption, so that what nodes do during the simulation is to optimistically assume that they can grant service for all of their pending requests, whereas in fact they can only be sure to be able to honor one such request. The consequence of this optimism is that, if *n*, concludes that it is not deadlocked, what this conclusion means is that there exists at each of the nodes in the wait graph an order according to which service should be granted by that node so that *n* will not deadlock. Of course that order may happen not to be followed and then *n* may deadlock in future global states.

In Algorithm *A_Detect_Deadlock*, node *n* maintains the following additional variables. The Boolean variable *notified*, initialized to *false*, is employed to indicate whether *n* has received at least one *notify*. Another variable is a counter, *granted*, (initialized to zero), to keep track of the number of *grant*’s *n* receives during the simulation. Two other counters, *in_dones* and *out_dones*, both initially equal to zero, indicate respectively the number of *grant_done* and *notify_done* messages received. The variables *in_parent* and *out_parent* are both initialized to *nil*. Node *n* detects that it is deadlocked if and only if *granted* $< x$, at the end.

**Algorithm A_Detect_Deadlock:**

Variables:

```
in_parent; = nil;
out_parent; = nil;
notified; = false;
granted; = 0;
in_dones; = 0;
out_dones; = 0.
```

**Listing 6.18**

Input:

```
msgi; = nil.
```

**Action if** $n_i \in N_0$:

```
notified; := true;
```
Send notify to all \( n_j \in \text{out}\_\text{wait} \).

**Listing 6.19**

**Input:**
\[ \text{msg}_i = \text{notify} \text{ such that } \text{origin}(\text{msg}_i) = (n_i, n_j). \]

**Action:**
- if \( \text{notified}_i \) then
  - Send \( \text{notify}\_\text{done} \) to \( n_j \)
- else
  - begin
    - \( \text{notified}_i := \text{true}; \)
    - \( \text{in}\_\text{parent}_i := n_j; \)
    - if \( |\text{out}\_\text{wait}_i| = 0 \) then
      - Send \( \text{grant} \) to all \( n_k \in \text{in}\_\text{wait}_i \)
    - else
      - Send \( \text{notify} \) to all \( n_k \in \text{out}\_\text{wait}_i \)
  - end.

**Listing 6.20**

**Input:**
\[ \text{msg}_i = \text{grant} \text{ such that } \text{origin}(\text{msg}_i) = (n_i, n_j). \]

**Action:**
- \( \text{granted}_i := \text{granted}_i + 1; \)
- if \( |\text{in}\_\text{wait}_i| = 0 \) then
  - Send \( \text{grant}\_\text{done} \) to \( n_j \)
- else
  - if \( \text{granted}_i \neq x_i \) then
    - Send \( \text{grant}\_\text{done} \) to \( n_j \)
  - else
    - begin
      - \( \text{out}\_\text{parent}_i := n_j; \)
      - Send \( \text{grant} \) to all \( n_k \in \text{in}\_\text{wait}_i \)
    - end.

**Listing 6.21**

**Input:**
\[ \text{msg}_i = \text{grant}\_\text{done}. \]

**Action:**
- \( \text{in}\_\text{dones}_i := \text{in}\_\text{dones}_i + 1; \)
- if \( \text{in}\_\text{dones}_i = |\text{in}\_\text{wait}_i| \) then
  - if \( \text{out}\_\text{parent}_i \neq \text{nil} \) then
    - Send \( \text{grant}\_\text{done} \) to \( \text{out}\_\text{parent}_i \)
  - else
    - Send \( \text{notify}\_\text{done} \) to \( \text{in}\_\text{parent}_i \)
Like several other asynchronous algorithms we have seen so far in the book (e.g., Algorithms A_PIF and A_Detect_Termination_D), this algorithm for deadlock detection by $n_1$ relies essentially on feedback information to achieve its purposes. Like those other algorithms, it maintains tree structures on the graph so that the feedbacks are sent only when appropriate.

In the case of Algorithm A_Detect_Deadlock, the pointers in_parent establish a tree that spans all the nodes that can be reached from $n_1$ in the wait graph. This tree is rooted at $n_1$ and its leaves are nodes for which the out_wait sets are empty. Its creation and eventual collapse are achieved by the pair (6.18) and (6.19), and by (6.22), respectively. In the same vein, for each of these nodes with empty out_wait sets, the pointers out_parent establish a tree that spans some of the nodes in the wait graph from which that node can be reached. Considered as a set of trees, they constitute a forest rooted at the nodes with empty out_wait sets spanning all the nodes in the wait graph from which at least one of the roots can be reached. The leaves of this forest are nodes whose grant messages sent during the simulation either never were the $x$th such message to reach their destinations or reached nodes with empty in_wait sets. This forest is created and collapses back onto its roots by means of (6.20) and (6.21), respectively.

It comes naturally from this discussion that the message complexity of Algorithm A_Detect_Deadlock is $O(m)$ while its time complexity is $O(n)$.

### 6.4 Exercises

1. **Show that Algorithm A_Self_Stabilize is still correct if the variables are restricted to 0,..., $V - 1$ for $V > n$.**

2. **For each of the asynchronous algorithms seen so far in the book (except for the templates), indicate the condition of global termination that allows the loop in Algorithm Task_t to be exited.**

3. **Give the details of node $n_1$'s participation in Algorithm A_Detect_Termination, as well as of the participation of all other nodes in the propagation of the terminate message.**

4. **Indicate how in Algorithm A_Detect_Termination the sending of reports to node $n_1$ can be simplified.**

5. **Discuss how to apply the technique of Section 6.2.2 to the cases in which $N_0$ does not contain one single element.**

6. **Repeat Exercise 3 for Algorithm A_Detect_Termination_D.**

7. **Show, in the context of Section 6.3, that the sets pending do not suffice to describe a graph that is necessarily between a wait graph and $G$.**

8. **Design an algorithm for global state recording, which, if initiated by one single node, is capable of informing that node of the global termination of the recording.**

   Show that Algorithm A_Self_Stabilize is still correct if the variables are restricted to 0,..., $V - 1$ for $V > n$. 

---

**Listing 6.22**

**Input:**

\[ \text{msg}_i = \text{notify}_\text{done}. \]

**Action:**

\[
\text{out\_dones}_i := |\text{out\_dones}_i| + 1;
\]

\[
\text{if out\_dones}_i = |\text{out\_wait}_i| \text{ then}
\]

\[
\text{if in\_parent}_i \neq \text{nil} \text{ then}
\]

Send notify done to in parent.
2. For each of the asynchronous algorithms seen so far in the book (except for the templates), indicate the condition of global termination that allows the loop in Algorithm Task_t to be exited. Give the details of node $n_t$’s participation in Algorithm A_Detect_Termination, as well as of the participation of all other nodes in the propagation of the terminate message. Indicate how in Algorithm A_Detect_Termination the sending of reports to node $n_t$ can be simplified.

3. Discuss how to apply the technique of Section 6.2.2 to the cases in which $N_0$ does not contain one single element. Repeat Exercise 3 for Algorithm A_Detect_Termination_D.

4. In Algorithm A_Detect_Termination the sending of reports to node $n_1$ can be simplified.

5. Show, in the context of Section 6.3, that the sets pending do not suffice to describe a graph that is necessarily "between" a wait graph and $G$.

6. Design an algorithm for global state recording, which, if initiated by one single node, is capable of informing that node of the global termination of the recording.

6.5 Bibliographic notes
The notion of self-stabilization was introduced by Dijkstra (1974), along with three algorithms on rings for which only considerably later proofs were provided (Dijkstra, 1986). Algorithm A_Self_Stabilize of Section 6.1 is based on one of the algorithms of Dijkstra (1974). For a survey of the investigations that this paper has spawned, the reader is referred to Schneider (1993). These investigations are quite broad in scope, ranging from a technique to prove self-stabilizing properties (Kessels, 1988) to applications to problems such as finding a minimum spanning tree (Chen, Yu, and Huang, 1991; Aggarwal and Kutten, 1993), computing on rings as in the original formulation by Dijkstra (1974) (Burns and Pachl, 1989; Flatebo and Datta, 1994), depth-first traversal (Huang and Chen, 1993; Collin and Dolev, 1994), leader election (Dolev and Israeli, 1992), coloring planar graphs (Ghosh and Karáta, 1993), finding maximal matchings (Hsu and Huang, 1992; Tel, 1994a), breadth-first numbering (Huang and Chen, 1992), establishing a local orientation on a ring (Israeli and Jalfon, 1991; 1993), and computing shortest distances (Tsai and Huang, 1994).


Section 6.2 is based on Huang (1989) for general computations and on Dijkstra and Scholten (1980) for diffusing computations. For an alternative account on the material in Dijkstra and Scholten (1980), the reader is referred to Bertsekas and Tsitsiklis (1989). Additional publications on termination detection include those by Chandrasekaran and Venkatesan (1990), Kavianpour and Bagherzadeh (1990), Ronn and Saikkonen (1990), Sheth and Dhamdhere (1991), Kumar (1992), Brzezinski, Hélary, and Raynal (1993), and Hélary and Raynal (1994).

Chapter 7: Graph Algorithms

Overview
The problems that we consider in this chapter are graph problems posed on $G$, similarly to what we did in Sections 4.2 and 4.3, in which we addressed the problems of graph connectivity and shortest distances, respectively. As in those sections, the aim here is to provide distributed algorithms in which all of $G$'s nodes participate in the solution based only on the partial knowledge of $G$'s structure that they have locally. However, our discussion in Section 3.2.2 should be recalled with special care throughout this chapter. Specifically, one alternative to the fully distributed approach we just mentioned is to elect a leader and have that leader obtain information on the entire structure of $G$. Having done this, the leader is then in position to solve the graph problem locally. As we remarked in that section, it takes $O(nm)$ messages and $O(n)$ time to concentrate all the relevant information in the leader, so that these two measures should be compared to the complexities of the fully distributed solution. But one must never lose sight of the possible impact of the resulting local time complexity (cf. Section 3.2) and of the implications of the nonconstant memory demand at the leader, in addition to the complexities associated with electing the leader in the first place.

We consider two graph problems in this chapter. The first problem is that of determining a minimum spanning tree on $G$. In addition to the role played by spanning trees in some of the problems we have studied so far, particularly in Sections 4.1 and 5.3, establishing a minimum spanning tree on $G$ is, as we remarked in Section 5.1, closely related to electing a leader in $G$, and then the relevance of the former problem is enlarged by its relation to all the situations in which having a leader is important. When a minimum spanning tree is sought with the purpose of electing a leader, then of course the alternative that we mentioned earlier of employing a leader to solve graph problems becomes meaningless. We deal with the minimum spanning tree problem in Section 7.1.

The other graph problem that we consider in this chapter is that of finding a maximum flow in a directed graph related to $G$ with certain characteristics. We address this problem in Section 7.2, where we present three asynchronous algorithms to solve it. What is interesting in our discussion in that section is that two of the algorithms that we discuss are originally conceived as synchronous algorithms. By employing the synchronization techniques we studied in Section 5.3, we may obtain a variety of corresponding asynchronous algorithms. Some of them are such that the resulting complexities of all the three asynchronous algorithms we consider are the same.

Sections 7.3 and 7.4 contain, respectively, exercises and bibliographic notes.

7.1 Minimum spanning trees

$G$ is in this section an undirected graph with FIFO edges. Our discussion is presented in three sections. Section 7.1.1 presents a statement of the problem, and Section 7.1.2 contains an asynchronous algorithm to solve it. Improvements leading to a reduced time complexity are given in Section 7.1.3.

7.1.1 The problem

As in Section 5.1, nodes in $N$ are assumed to have distinct identifications ($id(n)$ for node $n$) totally ordered by $<$. Associated with every edge $(n_i, n_j) \in E$ is a finite weight $w_{ij}$, known to both $n_i$ and $n_j$. The weight of a spanning tree is the sum of the weights of the $n - 1$ edges that constitute the tree. The minimum spanning tree problem asks that a spanning tree of minimum weight, called a minimum spanning tree, be found on $G$. Another related problem, that of finding any spanning tree on $G$, is clearly reducible to the problem of finding a minimum spanning tree, so that our discussion in this section applies to that problem as well. Although the problem of finding a spanning tree on $G$ (any one) is conceptually what we have needed in other occasions in this book (as in Sections 4.1 and 5.3), the more general problem has greater appeal for at least two reasons. The first reason is that edge weights can in some situations be used to model delays (or other related quantities) for message transmission over the edges, in which case a minimum spanning tree represents a tree of globally minimum transmission delay.
The other reason why considering the more general problem of determining a minimum spanning tree is more appealing is related to the use of such a tree as a first step in the election of a leader. A distributed algorithm to find a minimum spanning tree on $G$ can be built such that, at the end, every node has an indication of which edges incident to it are on the tree and which of these leads to the core of the tree, which is a single edge in the tree possessing properties that we describe later. Because only one core edge exists, the two nodes to which it is incident are natural candidates to be the leader. Clearly, under the usual assumption of totally ordered distinct identifications for all nodes, one of the two can be elected leader and the result broadcast over the tree with $O(n)$ message and time complexities. The core edge is only identified at the end of the algorithm, and may in principle be any edge, so the procedure we just described for leader election is only applicable to cases in which all of $G$'s nodes are candidates originally (cf. Section 5.1 for the appropriate terminology). If such is not the case, however, then the tree can still be employed as a basis to choose among the existing candidates according to the procedure we discussed in Section 5.1.

For simplicity, in this section we assume that all edge weights are distinct and totally ordered by $<$. If the particular connotation associated with edge weights poses difficulties with respect to this assumption, then the assumed existence of distinct identifications for all nodes can be used to break ties. Specifically, in such cases the weight of edge $(n_i, n_j)$ can be taken to be the pair $(n_i, \text{max}(id_i, id_j))$, and then all edge weights are totally ordered by $<$ in the lexicographic sense. Note that such a weight for edge $(n_i, n_j)$ can be computed easily by both $n_i$ and $n_j$ by simply sending their identifications to each other. All over $G$, this can be regarded as a first step in the computation of the minimum spanning tree. This first step requires $O(m)$ messages and $O(1)$ time, which, as we will see, does not add to the overall complexities of determining the minimum spanning tree.

A fragment of a minimum spanning tree is any subtree of the minimum spanning tree. An edge is said to be an outgoing edge of a fragment if one of the two nodes to which it is incident is in the fragment while the other is not. The distributed algorithms we study in this section to build a minimum spanning tree on $G$ are based on the following two properties (cf. Exercise 1).

i. If a fragment of a minimum spanning tree is enlarged by the addition of the fragment's minimum-weight outgoing edge, then the resulting subtree is also a fragment of the minimum spanning tree.

ii. If all edge weights are distinct, then $G$ has a unique minimum spanning tree.

Properties (i) and (ii) hint at the following basis for an algorithm to find a minimum spanning tree on $G$. Nodes in $N_0$ constitute single-node fragments initially. By property (i), these fragments can be enlarged independently of one another by simply absorbing nodes that are connected to the fragments by minimum-weight outgoing edges. Property (ii) ensures that the union of two fragments that grow to the point of sharing a node is also a fragment.

### 7.1.2 An algorithm

The algorithm we describe in this section employs properties (i) and (ii) along with the following rules for the creation of new fragments. The first rule is that every node in $N$ is initially a single-node fragment. This is achieved by having the nodes in $N_0$ broadcast a startup message by flooding over $G$ (similarly to the case of Algorithm $A_{Alg}(\text{Alpha})$ of Section 5.3.) with a message complexity of $O(m)$ and a time complexity of $O(n)$. Upon receiving a startup from every neighbor, a node initiates the algorithm as a single-node fragment. The second rule is that every fragment with at least two nodes has a special edge, called the core of the fragment, whose weight is taken to be the identification of the fragment. When the fragment is large enough to encompass all nodes (and then by properties (i) and (ii) it is the minimum spanning tree), its core is the tree's core, alluded to in the previous section.

The third overall rule regulates the process whereby fragments are combined to yield larger fragments. This combination is based on the level of each fragment, which is a nonnegative integer determined as follows. The level of a single-node fragment is zero. Now consider a
fragment at level \( \ell \geq 0 \), and let \( \ell' \) be the level of the fragment to which the fragment of level \( \ell \) is connected by its minimum-weight outgoing edge. If \( \ell = \ell' \) and the minimum-weight outgoing edges of both fragments are the same edge, then the two fragments are combined into a new fragment, whose level is set to \( \ell + 1 \) and whose core is the edge joining the former level-\( \ell \) fragments.

If \( \ell \neq \ell' \) or the two fragments' minimum-weight outgoing edges are not the same edge, then there are additional five cases to be considered. In two of the cases, \( \ell \neq \ell' \) and the two minimum-weight outgoing edges are the same. In these cases, the lower-level fragment is absorbed by the higher-level fragment and the resulting fragment inherits the higher level. In the remaining three cases, the two minimum-weight outgoing edges are not the same, and either \( \ell < \ell' \), or \( \ell = \ell' \), or \( \ell > \ell' \). In the case of \( \ell < \ell' \), the absorption of the level-\( \ell \) fragment by the level-\( \ell' \) fragment takes place just as we described earlier. If \( \ell \geq \ell' \), then the level-\( \ell \) fragment simply waits until the level of the other fragment has increased from \( \ell' \) enough for the combination to take place via one of the other possibilities.

Before we proceed, we should pause to investigate whether the waiting of fragments upon one another may ever lead to a deadlock. Specifically, the only situation one might be concerned about is that of a wait cycle comprising fragments, all of the same level, and such that the minimum-weight outgoing edge of every fragment leads to the next fragment in the cycle. By property (i), however, no such cycle may exist, as all the minimum-weight outgoing edges would have to be in the minimum spanning tree, which is impossible because they form a cycle.

Another property that may be investigated without any further details on how the algorithm functions is given by the following lemma.

**Lemma 7.1.**
The level of a fragment never exceeds \( \lfloor \log n \rfloor \).

**Proof:** For \( \ell > 0 \), a fragment of level \( \ell \) is only formed when two level-(\( \ell - 1 \)) fragments are such that their minimum-weight outgoing edges lead from one fragment to the other. An immediate inductive argument shows that a level-\( \ell \) fragment must then contain at least \( 2^\ell \) nodes (this holds for \( \ell = 0 \) as well), so \( n \geq 2^\ell \), thence the lemma.

Let us now provide the details of an algorithm to find a minimum spanning tree on \( G \) based on the overall strategy we just outlined. The algorithm is called Algorithm \( A_\text{Find_MST} \) ("MST" for Minimum Spanning Tree), and essentially proceeds repeatedly as follows, until the minimum spanning tree is found. First the minimum-weight outgoing edge of all fragments must be determined, then fragments must be combined with one another, and then (if the combination yielded a new, higher-level fragment) new fragment cores must be determined. During an execution of Algorithm \( A_\text{Find_MST} \), node \( n \) maintains a variable \( \text{state}_i \), which may be one of \text{find} or \text{found}, initially, \( \text{state}_i = \text{found} \), and along the-execution \( \text{state}_i \) switches back and forth between the two possibilities, indicating whether \( n \) is involved in the process of determining its fragment's minimum-weight outgoing edge (\( \text{state}_i = \text{find} \) or not (\( \text{state}_i = \text{found} \)). For each edge \((n_i, n_j) \in E\), \( n_i \) also maintains a variable which can be one of \text{on_tree}, \text{off_tree}, or \text{basic}, to indicate respectively whether the edge has been found by \( n_i \) to be an edge of the minimum spanning tree, not to be an edge of the minimum spanning tree, or still neither. Initially, this variable is set to \text{basic} for all \( n_j \in \text{Neig}_i \).

When a minimum-weight outgoing edge has been found for a fragment of level \( \ell \), a message \( \text{connect}(\ell) \) is sent over that edge. If such an edge is \((n_i, n_j) \) and \( n_i \) belongs to the level-\( \ell \) fragment, than such a message is sent by \( n_i \). There are two possibilities for the response that \( n_j \) gets from \( n_i \), whose fragment we take to be at level \( \ell' \). It may receive another \( \text{connect}(\ell) \), meaning that \( \ell = \ell' \) and \((n_i, n_j) \) is both fragments' minimum-weight outgoing edge, or it may happen that \( \ell < \ell' \). In the former case, the two fragments are joined into a level-(\( \ell + 1 \)) fragment whose core is \((n_i, n_j) \) and whose identification is the weight \( w_{ij} \). Nodes \( n_i \) and \( n_j \) are referred to as the "coordinators" of the new fragment, and their first task is to broadcast over the fragment the new level and new identification, as well as to direct all nodes in the...
fragment to begin a new search for a minimum-weight outgoing edge. The message that this broadcast carries is an *initiate*(ℓ + 1, w, find), where the find is the instruction for every node in the fragment to participate in looking for the fragment's minimum-weight outgoing edge. In the latter case, i.e., ℓ < ℓ', n's fragment absorbs n's fragment. In order to do this, n sends n either a message *initiate* (ℓ', w, find) or a message *initiate*(ℓ', w, found), where w is the identification of the fragment to which n belongs. This message is then broadcast by n over its own fragment to inform every node of their new fragment's level and identification. In addition, it prompts nodes to behave differently depending on whether a find or a found is in the message. If it is a find, then the nodes join in the search for the minimum-weight outgoing edge of the fragment they now belong to. If it is a found, then the nodes simply acquire information on their new fragment's level and identification. What remains to be explained on this interaction between n and n is the choice that n makes between attaching a find or a found to the initiate message that it sends. Node n attaches a find if state = find: it attaches a found if state = found. Sending a found in the initiate message is only correct if it can be argued that the weight of n's fragment's minimum-weight outgoing edge is strictly less than w, so that no edge outgoing from n's fragment could possibly be a candidate (because (n, n) is that fragment's minimum-weight outgoing edge). We provide this argument in what follows. The remaining cases cause n to wait for the level of n's fragment to increase from ℓ'.

So far we have seen that the coordinators of a newly formed fragment broadcast initiate messages with a find parameter over the edges of the new fragment. This broadcast is meant to inform all the nodes in the fragment that the fragment has a new level and a new identification. It also carries a find parameter that directs the nodes to engage in seeking the minimum-weight outgoing edge of the new fragment. A node n that is reached by an initiate message with a find parameter sets state to find and participates in locating the fragment's minimum-weight outgoing edge. When n's participation in this process is finished, then state, is reset to found. If, on the other hand, the initiate message carries a found parameter, then its effect upon n is simply the fragment level and identification update. The broadcast of an initiate message may go beyond the boundaries of the fragment if a node n that it has reached receives a connect message from another fragment whose level is strictly less than the level being carried by the initiate message. The broadcast is then propagated through that fragment as well, representing its absorption by the higher-level fragment. The initiate messages that n propagates into the lower-level fragment carry either a find or a found parameter, depending on whether state = find or state = found. Let us now discuss in detail the process whereby the minimum-weight outgoing edge of a fragment is found. If the fragment has level zero, and therefore comprises one single node, then that node simply inspects the edges that are incident to it and sends a connect(0) message over the edge having minimum weight. In addition, if that node is n, then state is set to found. If the fragment's level is strictly positive, then it must rely on the initiate message broadcast by its coordinators to have all the nodes participate in the process. After receiving an initiate(ℓ, w, find) and setting state to find, node n considers all edges (n, n) for which

\[ \text{state}_{i}^{j} = \text{basic} \]

in increasing order of weights. On each edge that it considers, n sends a test(ℓ, w) and waits to receive either a reject message or an accept message. If ℓ' is the level of the fragment to which n belongs and w' that fragment's identification (or at least n's view of that level and that identification, which may already have changed), then the reject is sent by n, after it sets \( \text{state}_{i}^{j} \) to off_tree, if w = w' (in this case, n and n are in the same fragment and the edge between them cannot possibly be on the minimum spanning tree). If w ≠ w' and ℓ ≥ ℓ, then n sends n an accept. If w ≠ w' and ℓ < ℓ, then n is not in position to send any response immediately and waits to do so until its level has increased to be at least equal to ℓ (at which time it must also re-evaluate the relation between w and w', as the latter may have changed along with ℓ'). An accept received from n makes n stop the search. A reject that it receives from n causes it to set

\( \text{state}_{i}^{j} \) to off_tree. When n receives an accept from a neighbor n, the edge (n,
higher-level fragments upon lower-level ones. For all send its is used by The weight of fragment to which it belongs are denoted respectively by indicates whether a employed the following variables at node We now turn to the presentation of the algorithm's actions. In addition to the variables that we have already introduced during our preceding discussion, Algorithm also initiates waves over edge. If such an edge is \((n, n_i)\), then \(n_i\) sets \(\text{state}^j_i = \text{on\_tree}\) as it sends the message. As in various occasions so far in the book, interpreting Algorithm as propagating waves over \(G\) can be very helpful in building some intuitive understanding on how it works. What happens in this case is that a fragment's core propagates a wave of messages over \(\text{on\_tree}\) edges. This wave collapses back with report messages onto the core, and then a new fragment is formed after the change\_core and connect messages have played their roles. The \(\text{initiate}\) waves may occasionally "leak" from the fragment when neighboring fragments are absorbed.

We now turn to the presentation of the algorithm's actions. In addition to the variables that we have already introduced during our preceding discussion, Algorithm also employs the following variables at node \(n_i\). The Boolean \(\text{initiated}\), initially equal to \(\text{false}\), is used to indicate whether \(n_i \in N_n\). For all \(n_i \in \text{Neig}\), the Boolean \(\text{go}^j_i\), equal to \(\text{false}\) initially, indicates whether a \(\text{startup}\) has been received from \(n_i\). At \(n_i\), the level and identification of the fragment to which it belongs are denoted respectively by \(\text{level}\) (set to zero initially) and \(\text{frag}\). The weight of \(\text{best\_edge}\), is stored in \(\text{best\_weight}\). The counter \(\text{expected}\), initialized to zero, is used by \(n_i\) to indicate the number of messages it must receive before being in position to send its report message. An additional group of variables is used to control the wait of higher-level fragments upon lower-level ones. For all \(n_i \in \text{Neig}\), these are the Booleans \(\text{got\_connect}^j_i\) and \(\text{got\_test}^j_i\), both initialized to \(\text{false}\), and also...
The two Booleans are used, respectively, to indicate that a connect or a test has been received from a higher-level fragment and cannot therefore be replied to at once. When $test_{\text{frag}}_i$ = true, then $test_{\text{level}}_i$ and $test_{\text{frag}}_i$ store respectively the level and the identification that the test message carried.

**Algorithm A_Find_MST:**

**Variables:**
- initiated = false;
- $go_i^j$ = false for all $n_i \in \text{Neig}$;
- state = found;
- level $i$ = basic for all $n_i \in \text{Neig}$;
- frag;
- best_edge;
- best_weight;
- to_core;
- expected $i$ = 0;
- $got_{\text{connect}}_i^j$ = false for all $n_i \in \text{Neig}$;
- $got_{\text{test}}_i^j$ = false for all $n_i \in \text{Neig}$;
- $test_{\text{level}}_i^j$ for all $n_i \in \text{Neig}$;
- $test_{\text{frag}}_i^j$ for all $n_i \in \text{Neig}$.

**Listing 7.1**

**Input:**
- $msg_i = \text{nil}$.

**Action if** $n_i \in N_0$:
- initiated := true;
- Send startup to all $n_i \in \text{Neig}$.

**Listing 7.2**

**Input:**
- $msg_i = \text{startup}$ such that $\text{origin}(msg) = (n_i, n_j)$
Action:
  if not initiated, then
  begin
    initiated, := true;
    Send startup to all \( n \in \text{Neig}_i \);
  end;

  \( go^j_i \) := true;

  if \( go^j_i \) for all \( n \in \text{Neig}_i \) then
  begin
    Let \((n_i, n_k)\) be such that \( w_{ik} \leq w_{il} \) for all \( n_l \in \text{Neig}_i \);
    \( \text{state}^k_i \) := on_tree;
    Send connect(level) to \( n_j \);
  end.

Listing 7.3

Input:
\( msg_i = \text{connect}(l) \) such that \( \text{origin}(msg_i) = (n_i, n_j) \).
Action:
  if \( l < \text{level}_i \) then
  begin
    \( \text{state}^j_i \) := on_tree;
    Send initiate(level, frag, state) to \( n_j \);
    if \( \text{state}_i = \text{find} \) then
      expected, := expected, + 1
    end
  else

    if \( \text{state}^j_i = \text{basic} \) then
      \( \text{got_connect}^j_i \) := true
    else
      Send initiate(level, + 1, w, find) to \( n_j \).

Listing 7.4

Input:
\( msg_i = \text{initiate}(l, w, st) \) such that \( \text{origin}(msg_i) = (n_i, n_j) \).
Action:
  level, := \( l \);
  frag, := \( w \);
  for all \( n \in \text{Neig}_i \) such that \( \text{got_test}^k_i \) do
if \( \text{test}_i \leq \text{level} \), then

if \( \text{test}_i \neq \text{frag} \), then
  Send accept to \( n_k \)
else
  begin
    if \( \text{state}_i = \text{basic} \) then
      \( \text{state}_i := \text{off}_{\text{tree}}; \)
    Send reject to \( n_k \)
  end;
\( \text{state}_i := \text{st}; \)
\( \text{to}_{\text{core}} := (n_i, n_j); \)
\( \text{best}_\text{weight} := \infty; \)

for all \( n_k \in \text{Neig} \setminus \{n_i\} \) such that \( \text{state}_i = \text{on}_{\text{tree}} \) do
begin
  Send \text{initiate}(\text{level}_i, \text{frag}_i, \text{state}_i) to \( n_k \);
  if \( \text{state}_i = \text{find} \) then
    \( \text{expected}_i := \text{expected}_i + 1 \)
  end;
if \( \text{state}_i = \text{find} \) then
  if \( n_k \in \text{Neig} \) exists such that \( \text{state}_i = \text{basic} \) then
  begin
    Let \( B \subset \text{Neig} \) be such that \( \text{state}_i = \text{basic} \) for all \( n_k \in B \);
    Let \( n_k \in B \) be such that \( w_{ik} \leq w_{il} \) for all \( n_l \in B \);
    Send \text{test}(\text{level}_i, \text{frag}_i) to \( n_k \)
  end
else
  if \( \text{expected}_i = 0 \) then
  begin
    \( \text{state}_i := \text{found}; \)
    Send \text{report}(\text{best}_\text{weight}) on \( \text{to}_{\text{core}} \); end.

---

**Listing 7.5**

**Input:**
\( \text{msg}_i = \text{test}(\ell,w) \) such that \( \text{origin}(\text{msg}_i) = (n_i, n_j) \).

**Action:**
if \( \ell > \text{level} \), then
begin
  \( \text{got}_{\text{test}_i} := \text{true}; \)
Listing 7.6
Input: 
\[ msg_i = \text{accept such that } \text{origin}(msg_i) = (n_i, n_j). \]
Action:
\[ \text{if } w_{ij} < \text{best_weight}_i \text{ then} \]
\[ \quad \begin{align*}
\text{best_weight}_i & := w_{ij}; \\
\text{best_edge}_i & := (n_i, n_j)
\end{align*} \]
\[ \text{end;} \]
\[ \text{if } \text{expected}_i = 0 \text{ then} \]
\[ \quad \begin{align*}
\text{state}_i & := \text{found}; \\
\text{Send } \text{report } (\text{best_weight}) \text{ on } \text{to_core}_i
\end{align*} \]
\[ \text{end.} \]

Listing 7.7
Input:
\[ msg_i = \text{reject such that } \text{origin}(msg_i) = (n_i, n_j). \]
Action:
\[ \text{if } \text{state}_i = \text{basic then} \]
\[ \quad \text{state}_i := \text{off_tree}; \]
\[ \text{if } n_k \in \text{Neig}_i \text{ exists such that } \text{state}_k = \text{basic} \text{ then} \]
\[ \quad \begin{align*}
\text{state}_i & := \text{off_tree}; \\
state_{i}^{j} & := \text{basic}
\end{align*} \]
Let $B \subseteq \text{Neig}_i$ be such that $\forall n_k \in B$; $\forall n_i \in B$; $w_{nk} \leq w_i$; $\forall n_i \in B$; $\forall n_k \in B$; $\forall n_l \in B$; $w_{ik} \leq w_{il}$.

Send test(level, frag) to $n_k$.

**Listing 7.8**

**Input:**
$\forall n_i \in \text{Neig}_i$: $\exists n_k \in B$ such that $(n_i, n_k) \in \text{Neig}_i$.

**Action when** $(n_i, n_j) \neq \text{to-core}, or \ state_i = \text{found}:$ 
if $(n_i, n_j) \neq \text{to-core}, then$

begin
expected_i := expected_i - 1;
if $w < \text{best_weight}_i$ then
begin
$\text{best_weight}_i := w_i$;
$\text{best_edge}_i := (n_i, n_j)$
end;
if $\text{expected}_i = 0$ then
begin
state_i := $\text{found}$;
Send report(best_weight) on to_core_i;
end
end

else
if $w > \text{best_weight}_i$ then
begin
Let $n_k \in \text{Neig}_i$ be such that $(n_i, n_k) = \text{best_edge}_i$;
if $\text{state}_i = \text{on_tree}$ then
Send change_core on best_edge_i;
else
begin
Send connect(level) on best_edge_i;
end
end
end
end.
Listing 7.9

Input:
\[ msg = \text{change\_core} \text{ such that origin}(msg_i) = (n_i, n_i). \]

Action:
Let \( n_i \in Neig \) be such that \((n_i, n_i) = \text{best\_edge}_i; \)
\[
\text{if} \quad \text{state}_i^k = \text{on\_tree} \quad \text{then}
\]
\[
\quad \text{Send} \quad \text{change\_core} \quad \text{on} \quad \text{best\_edge};
\]
\[
\text{else}
\]
\[
\quad \text{begin}
\]
\[
\quad \quad \text{Send} \quad \text{connect}(\text{level}) \quad \text{on} \quad \text{best\_edge};
\]
\[
\quad \quad \text{if} \quad \text{got\_connect}_i^k \quad \text{then}
\]
\[
\quad \quad \quad \text{Send} \quad \text{initiate}(\text{level} + 1, w_{ik}, \text{find}) \quad \text{on} \quad \text{best\_edge};
\]
\[
\quad \text{state}_i^k := \text{on\_tree}
\]
end.

Actions (7.1) through (7.9) implement the overall strategy we described in detail earlier to find a minimum spanning tree on \( G \). These actions, the reader must have noticed, account for far more complex a behavior than that of any of the algorithms we have seen (or will see) in other chapters. Although a complete proof of correctness cannot be offered within the scope of this book, we now pause momentarily to offer some more detailed comments on each of the actions, so that the reader may have additional guidance in studying them.

Actions (7.1) and (7.2) are the standard initial actions so that all nodes can begin participating in the algorithm after the initial flood of startup messages. A node's initial participation consists of sending a connect(0) message over the minimum-weight edge that is incident to it. Upon receiving a connect(\( \ell \)) message from \( n_j \) in (7.3), \( n_i \) either immediately absorbs the originating fragment (if \( \ell < \text{level}_i \)), or it recognizes that this connect is the response to a connect that it sent previously on the same edge, and therefore the two fragments must be merged into another of higher level (if \( \ell \geq \text{level}_i \) and \( \neq \text{basic} \)). If \( \ell \geq \text{level}_i \) and \( \text{state}_i^k \neq \text{basic} \), then this must be a connect from a higher-level fragment and must not be replied to immediately.

The receipt of an initiate message by \( n_i \) in (7.4) first causes the node to update its fragment level and identification and then to reply to any of its neighbors that may have sent a test message in the past with a level higher than its own. It then forwards the initiate message on all the other on_tree edges that are incident to it and, if \( \text{state}_i = \text{find} \), begins the search for its minimum-weight outgoing edge by means of test messages, if basic edges exist that are incident to it (otherwise, it may be in position to send its report).

When node \( n_i \) receives a test(\( \ell, w \)) message from \( n_j \) in (7.5) and \( \ell > \text{level}_i \), then it cannot reply immediately and saves both \( \ell \) and \( w \) for later consideration when its own level increases in (7.4). If \( \ell \leq \text{level}_i \), then either an accept gets sent to \( n_i \) (if \( w \neq \text{frag} \)) or a reject gets sent (otherwise).

The receipt of an accept by \( n_i \), in (7.6) may cause best_edge, to be updated (along with best_weight), and may in addition signal to \( n_i \) that it may send its report message. If a reject is received in (7.7) and there are additional basic edges incident to \( n_i \), then the node
continues its probing with test messages; if no such edges are left, then $n_1$ checks whether its report may be sent.

Upon receiving a report($w$) message in (7.8), there are two possibilities for $n_1$. The first possibility is that the message is received on edge to_core, in which case $n_1$ must be a coordinator of the fragment and has to decide on which side of the core the fragment's minimum-weight outgoing edge lies. If that edge is to be found on its own side (i.e., if $w > \text{best_weight}$), then either it sends a change_core or a connect on best_edge, the former if best_edge is an on_tree edge, the latter otherwise (and then the fragment's minimum-weight outgoing edge is incident to $n_1$, thence the connect that it sends). The second possibility is that of $(n_1, n_2) \neq \text{to_core}$, in which case $n_1$ checks whether it is time for its own report to be sent.

It is important to notice, in (7.8), that the action is only executed upon receipt of the report on $(n_1, n_2)$ when $(n_1, n_2) \neq \text{to_core}$, or state = found. This ensures that a report arriving on the core $((n_1, n_2) = \text{to_core})$ is only acted upon when $n_1$ has already identified the least weight on its side of the fragment (state = found) and may therefore decide on the fragment's minimum-weight outgoing edge. (Associating Boolean conditions to actions can also be an approach to delaying the receipt of a message that cannot be replied to immediately, as in (7.3) and (7.5); however, this can only be done in the presence of edges that are not FIFO—cf. Exercise 2.)

Action (7.9) corresponds to the receipt by $n_1$ of a change_core message, which is either forwarded on best_edge, or causes a connect to be sent on that edge, depending on whether the fragment's minimum outgoing edge is incident to $n_1$ just as in the case of (7.8). When sending a connect in either (7.8) or (7.9), $n_1$ may also have to send an initiate after it, if in (7.3) a connect was received that could not be replied to immediately.

The algorithm's termination is detected by each coordinator $n_1$ upon receiving a report($w$) message on the core when state = found such that $w = \text{best_weight} = \infty$. After the minimum spanning tree has been found on $G$, at every node the on_tree edges indicate which of the edges incident to it are on the tree, while the to_core edge indicates which of the on_tree edges leads to the tree's core.

Next we present Algorithm A_Find_MST's complexities.

**Theorem 7.2.**

Algorithm A_Find_MST has a message complexity of $O(mn \log n)$ and a time complexity of $O(n \log n)$. In addition, the algorithm's bit complexity is $O(m + n \log n) (\log W + \log \log n)$, where $W \geq |w_i|$ for all $(n_1, n_2) \in E$.

**Proof:** Let $l$ and $w$ denote respectively a generic fragment level and edge weight. A node can never send more than one reject message on the same edge in the same direction. In addition, to each such message there corresponds a test($l$, $w$) message, therefore accounting for $O(m)$ messages and, by Lemma 7.1, $O(m(\log W + \log \log n))$ bits. At each level, a node can receive at most one initiate ($l$, $w$, $st$) and one accept, and it can send no more than one test($l$, $w$) resulting in an accept, one report ($w$), and one change_core or connect($l$), where $st$ is one of find or found and requires a constant number of bits to be expressed. By Lemma 7.1, we have another $O(n \log n)$ messages and $O(n \log n(\log W + \log \log n))$ bits, which, added to what we already have, yields the algorithm's message and bit complexities.

The algorithm's time complexity follows directly from Lemma 7.1 and from the observation that, for each level, the propagation of messages within a fragment takes no more than $O(n)$ time.

It should be noted that the initial complexities for determining edge weights (if not distinct originally, in which case node identifications must be used) and for exchanging the startup messages do not add to the complexities we have determined.

We finalize the section by returning to some issues raised earlier in the book. The first issue is that of electing a leader once the minimum spanning tree has been found. As we observed earlier in Section 7.1.1, the final coordinators (nodes to which the tree's core is incident) may elect a leader in $O(n)$ time and with $O(n)$ messages. The resulting complexities for the leader election (including those of finding the minimum spanning tree) are then the same as those given by Theorem 7.2. When compared with the $O(n \log n)$-message, $O(n)$ time procedure for
leader election described in Section 5.1, the new approach has a better message complexity, but its time complexity turns out to be somewhat worse.

The second issue is that of the complexities to initialize Synchronizer Beta in Section 5.3.1. The reader should recognize quickly that Theorem 7.2, together with the observation we just made on the election of a leader on the tree, provides the values of $\text{Messages}_s$ and $\text{Time}_s$ for Synchronizer Beta.

### 7.1.3 Further improvements

Although it can be argued that the $O(m + n \log n)$ message complexity is the best one can hope for when finding a minimum spanning tree on $G$, reducing the time complexity from the $O(n \log n)$ of Algorithm $A_{\text{Find}_MST}$ has been the subject of investigations, aiming at bringing it down to $O(n)$. We do not in this section aim at conveying the details of how this improved time complexity can be achieved, but rather point at some of the inessential sources of time complexity in Algorithm $A_{\text{Find}_MST}$ and at some possible improvements.

In order to identify the reason for the excessive time complexity of Algorithm $A_{\text{Find}_MST}$, we must look at the proof of Lemma 7.1, where we argue that, for $t \geq 0$, a level-$t$ fragment has at least $2^t$ nodes. Although for $t = 0$ this number is exactly $2^t$ (level-0 fragments comprise exactly one node), for $t > 0$ a level-$t$ fragment may include a lot more than the minimum $2^t$ nodes. Because the algorithm is such that higher-level fragments wait for lower-level fragments to have their levels increased before they can be merged, a level-$(t+1)$ fragment that happens to be waiting for such an oversized level-$t$ fragment may have to wait for as long as $O(n)$ time before the merge (this is what is argued in the proof of Theorem 7.2 as far as the time complexity is concerned).

So the attempts at improving the algorithm's time complexity concentrate on relating a fragment's size to its level more tightly. One such attempt is, for example, to reduce the number of fragments by an $O(\log n)$ factor within $O(n)$ time. As a consequence, the reduction from the initial $n$ fragments to the final single fragment representing the minimum spanning tree can be achieved in as many $O(n)$-time portions as it takes at a rate of $O(\log n)$ per portion. Employing the usual notation $\log^* k$ to denote the number of times log has to be applied to reduce $k > 1$ to a number no greater than one, we see that the number of $O(n)$-time portions that we need is $\log^* n$. The time complexity of the resulting algorithm is then $O(n \log^* n)$, and its message complexity can be shown to remain the same as that of Algorithm $A_{\text{Find}_MST}$.

### 7.2 Maximum flows in networks

In all of Section 7.2, $G$ is an undirected graph with two distinguished nodes, called a source (which we assume to be $n_1$) and a sink (which we assume to be $n_n$). G is an undirected graph, in conformity with our practice in this book that only in such graphs may communication between neighbors flow in both directions. However, the denominations of $n_1$ and $n_n$, respectively as source and sink are only meaningful when we consider a directed variation of $G$, denoted by $G^d$, which is the graph on which the problem that we deal with is posed. The reason for employing the two graphs is that we want to be able to state the problem properly and yet, during the execution of the algorithms that we shall investigate, be able to have messages sent between neighbors in both directions.

The directed graph $G^d$ is obtained from $G$ by associating a direction with each of $G$'s edges. These directions are such that $n_1$ must not have any edge directed toward itself (thence its denomination as a source) and $n_n$ must not have any edge directed away from itself (thence its denomination as a sink). In addition, $n_1$ must be the only source in $G^d$ and $n_n$ the only sink, which implies that all the other nodes must in $G^d$ lie on a directed path from $n_1$ to $n_n$. As a side remark, the reader should notice that, in Section 6.3.2, we were faced with the same notational issue of being able to refer to $G$ as an undirected graph and at the same time to another directed graph defined as a function of $G$. A similar situation will occur once again in the book, specifically in Sections 8.3 and 8.4.

The problem that we study in the next three sections is the problem of computing a maximum flow in $G^d$, which in this context is referred to as a "network," although we refrain from employing this denomination any further in the book, lest there may be confusion with
Analogously, if $n_i = n_j$, then property (iii) does not hold, because

$$\sum_{(n_i \rightarrow n_j) \in E^d} f(n_i, n_j) \geq 0$$

and

$$\sum_{(n_i \rightarrow n_j) \in E^d} f(n_j, n_i) = 0.$$ 

Analogously, if $n_i = n_m$, then

$$\sum_{(n_j \rightarrow n_i) \in E^d} f(n_j, n_i) = 0, \quad \sum_{(n_j \rightarrow n_i) \in E^d} f(n_i, n_j) = 0.$$
\[ \sum_{(i_n, n_i) \in E^d} f(n_i, n_i) \geq 0 \]

and

\[ \sum_{(i_n, n_i) \in E^d} f(n_i, n_i) = 0. \]

The value of a flow \( f \), denoted by \( F \), is given by the summation in either of the two previous inequalities, that is,

\[ F = \sum_{(n_1 \rightarrow n_2) \in E^d} f(n_1, n_2) = \sum_{(n_j \rightarrow n_n) \in E^d} f(n_j, n_n) \]

(cf. Exercise 3). The maximum-flow problem asks for a flow \( f \) of maximum value. For \( n_i, n_j \in N \), the residual capacity of the ordered pair \( (n_i, n_j) \) given a flow \( f \) is

\[ c_f(n_i, n_j) = c(n_i, n_j) - f(n_i, n_j), \]

being therefore equal to zero if \( (n_i, n_j) \in E \). Readily, \( c(n_i, n_j) \geq 0 \) if \( (n_i \rightarrow n_j) \in E^d \), whereas \( c \) \( f(n_i, n_j) \) (\( n_i \rightarrow n_j \)) \( \in E^d \). The residual network of \( G \) given \( f \) is the directed graph \( G_f = (N, E_f) \) and is such that \( (n_i \rightarrow n_j) \in E_f \) if and only if \( (n_i, n_j) \in E \) and \( c(n_i, n_j) > 0 \). Clearly, if \( (n_i \rightarrow n_j) \in E^d \), then both \( (n_i, n_j) \) and \( (n_j, n_i) \) may be members of \( E_i \), so long as \( f(n_i, n_j) < c(n_i, n_j) \) and \( f(n_j, n_i) \geq 0 \) (these are, respectively, the conditions for each of the memberships in \( E_i \)). A directed path from \( n_i \) to \( n_j \) in \( G \) is called an augmenting path. The intuitive support for this denomination is that, along such a path, the residual capacity of \( (n_i \rightarrow n_j) \in E_i \) can be decreased by either increasing \( f(n_i, n_j) \) if \( (n_i \rightarrow n_j) \in E^d \) or decreasing \( f(n_j, n_i) \) for \( (n_i \rightarrow n_j) \in E^d \).

When \( f \) does not satisfy property (iii), but rather the weaker property that

\[ \sum_{n_i \in N} f(n_i, n_j) \leq 0 \]

for all \( n_i \in N - \{n_i, n_j\} \) then it is called a preflow instead of a flow. In this case, there exists an excess flow coming into \( n_i \), denoted by \( e_i(n) \) and given by

\[ e_i(n) = -\sum_{n_j \in N} f(n_i, n_j). \]

The next two sections are devoted to the presentation of three distributed algorithms for the maximum-flow problem. Two of these algorithms are synchronous and appear in Section 7.2.2. The other algorithm is asynchronous, and is presented in Section 7.2.3. The first of the synchronous algorithms is based on the concepts of residual networks and augmenting paths, and is called Algorithm \( S_{\text{Find}_\text{Max}_\text{Flow}} \). The other synchronous algorithm and the asynchronous algorithm are both based on the concept of preflows. These two algorithms are considerably simpler than Algorithm \( S_{\text{Find}_\text{Max}_\text{Flow}} \), and for this reason are not presented with all the details as that one is.

Algorithm \( S_{\text{Find}_\text{Max}_\text{Flow}} \) and the preflow-based synchronous algorithm can both be shown to have the same message and time complexities, being respectively of \( O(n^3) \) and \( O(n^2) \). The asynchronous algorithms that result from applying Synchronizer \( \Gamma \) to Algorithm \( S_{\text{Find}_\text{Max}_\text{Flow}} \) and to the other synchronous algorithm, following our discussion in Section 5.3.1, both have message complexity and time complexity, for \( 2 \leq k < n \), respectively of \( O(km^2) \) and \( O(n^2 \log n / \log k) \). If Synchronizer \( \Gamma \) is used instead, then the resulting asynchronous algorithms have message complexity of \( O(n^2m) \) and time complexity of \( O(n^2) \). Interestingly, these are the complexities that the preflow-based asynchronous algorithm has been shown to have as well.

7.2.2 Two synchronous algorithms

The essence of Algorithm \( S_{\text{Find}_\text{Max}_\text{Flow}} \) is the following. It proceeds in iterations, and at each iteration a "layered" residual network (to be explained shortly) is built. A maximal flow is then found on this network and then added to the cumulative flow that is maintained throughout the iterations. A flow in the layered residual network is said to be maximal when it is equal to the residual capacity of at least one edge on every \( n_i \)-to-\( n_j \) path. When a layered residual network with at least one augmenting path can no longer be found, the flow is maximum and the algorithm terminates.
The layered residual network is built at each iteration as follows. Let \( f \) be the cumulative flow obtained at the end of the previous iteration (the initial flow, for the first iteration). The source \( n_1 \) is included in the first layer and a process similar to the breadth-first numbering discussed in Section 3.4 is started to determine the subsequent layers. For \( l > 1 \), the \( l \)th layer contains every node \( n_i \) that is not in any of the previous \( l - 1 \) layers and such that there exists a node \( n_j \) in the \( l - 1 \)st layer such that \( c(n_i, n_j) > 0 \). The synchronous algorithm to build the layered residual network is then very simple. For \( l > 1 \) and \( \sigma \geq 0 \) to indicate the pulses within each iteration, the \( l \)th layer is determined at pulse \( \sigma = l - 2 \) as follows. Those nodes \( n_i \) belonging to the \( l - 1 \)st layer send a message to their neighbors \( n_i \) such that \( c(n_i, n_j) > 0 \). In the next pulse (i.e., \( \sigma = l - 1 \)), \( n_i \) replies positively or not at all to \( n_j \), depending on whether it had already been included in a layer at any of the previous pulses.

Once the layered residual network has been constructed based on a flow \( f \), a maximal flow on it is determined by a process that is started at \( n_1 \) by assigning to each \( (n_i \rightarrow n_j) \in E \) a flow equal to \( c(n_i, n_j) \), thereby providing \( n_i \) with a positive excess flow. This process continues on to the succeeding layers, and along the way the excess flow at the nodes is either pushed to the next layer or returned to the previous one (and then possibly re-routed through other edges). Termination occurs when no node can take any additional flow. The synchronous algorithm to find a maximal flow on the layered residual network works by sending flow between neighbors in the form of messages. Whenever flow is received at \( n_i \) from \( n_j \) on edge \( (n_i, n_j) \) such that \( (n_i \rightarrow n_j) \in E \), the amount of flow received is pushed onto a stack along with a pointer to its sender, \( n_j \). At each pulse, a node \( n_i \) may receive flow from \( n_j \) on edge \( (n_i, n_j) \) such that \( (n_i \rightarrow n_j) \in E \) or such that \( (n_j \rightarrow n_i) \in E \) (this is returned flow). At the beginning of the next pulse, all the flow \( n_i \) received is either sent to the succeeding layer, if at all possible, or returned to the previous one, in this case by popping the amount of flow to be returned and its destination off the stack. In case no more flow can be sent to the next layer, \( n_i \) informs its neighbors in the preceding layer that it is "blocked," so no further attempts will be made to send flow to it in the remainder of the iteration.

Algorithm \( S\_Find\_Max\_Flow \) proceeds in iterations \( k = 1, \ldots, K \), where \( K \) is initially viewed as being equal to infinity and is set to its correct value upon detection by the nodes that the current iteration is the last one. Incidentally, the detection of termination in this case is, like for the other synchronous algorithms we have seen, essentially a matter of counting pulses as they elapse. However, as we mentioned in Section 6.2, in this case such a strategy is supported by the nontrivial arguments (which we do not reproduce here) that lead to the algorithm's time complexity.

For \( 1 \leq k < K \), the \( k \)th iteration comprises two phases, each no more than \( 2n \) pulses long. The first phase of an iteration is used to find the layered residual network, while the second phase is used to find a maximal flow on that network, so for the \( k \)th iteration only the first phase is needed. The value of \( K \), however, can only be known after a first phase in which \( n_1 \) could not be reached during the construction of the layered residual network has occurred.

Intuitively, Algorithm \( S\_Find\_Max\_Flow \) proceeds through the propagation of synchronous waves emanating from \( n_1 \). During the first phase of an iteration, such a wave expands from \( n_1 \) to construct the layered residual network, with feedback information sent to \( n_1 \) when the network is constructed. During the second phase, the wave that \( n_1 \) initiates pushes flow onward on the layered residual network, with occasional "ripples" of returned flow in the opposite direction that may in turn be sent onward again.

The following are the messages employed in Algorithm \( S\_Find\_Max\_Flow \). A message layer is employed to build the layered residual network. It is propagated, starting at \( n_1 \), on edges that have positive residual capacity (except the edge, if any, on which it was received) when received for the first time in an iteration. Every layer message, if belonging to the first group of such messages to be received in the current iteration, is replied to with an ack. This propagation of layer messages, as well as ack’s, accounts for at most the first \( n \) pulses of an iteration. Additional \( n \) pulses (at most) are employed for a success message to be sent by \( n_1 \) toward \( n_1 \) if it is reached by the layer messages. If within \( 2n \) pulses of the beginning of an iteration \( n_1 \) does not receive a success, then it may conclude that the layer messages did not reach \( n_1 \) and therefore may set \( K \) to the number of the current iteration to terminate the algorithm. The second phase employs flow(\( x \)) and block messages, respectively to ship an amount \( x \) of flow and to signal that the sender of the message should not be sent any more flow during the iteration.
Node \( n \) employs the following variables. A stack, initialized to nil, is employed for \( n \) to store the flow shipments it receives, and their origins, for later return if the need arises. The excess flow at \( n \) is stored in the variable excess, initially equal to zero. A Boolean reached, initially equal to false, indicates whether during the current iteration \( n \) has already been reached by a layer message. As in previous occasions, the node from which \( n \) receives layer for the first time is pointed to by parent, initially set to nil (if a layer is received from more than one neighbor at the same pulse, then the choice of which neighbor parent is to point to is arbitrary). For all \( n_i \in Neig \), the following variables are used. The Booleans

\[
\text{in}_{\text{previous}}_{\text{layer}}^j, \text{in}_{\text{next}}_{\text{layer}}^i, \text{residue}^i, \text{flow}^i,
\]

respectively to indicate, for each iteration, whether \( n \) is in the previous layer or in the next layer of the layered residual network with respect to \( n \). The variables

\[
\text{flow}^i, \text{residue}^i, \text{blocking}^i, \text{returned}^i
\]

give, respectively, the value of the current flow and current residual capacity of the ordered pair \((n, n_i)\). They are both initialized to zero, unless \((n, n_i) \in E\), in which case \text{residue}^i \text{is initialized to} \( c(n, n_i) \). The Booleans

\[
\text{in}_{\text{previous}}_{\text{layer}}^j, \text{in}_{\text{next}}_{\text{layer}}^i, \text{blocking}^i, \text{returned}^i
\]

are used to signal whether more flow can be sent to \( n \) during the current iteration and whether flow has been returned to \( n \) during the current iteration. Finally, node \( n \) employs an auxiliary variable \( y \).

The initial values we have given for the variables are employed either at the beginning of the algorithm, and they appear when the variables are first listed, or at the beginning of each iteration. Variables whose initial values are used only once at the beginning of the algorithm are the variables related to flows and capacities (these are the excess, flow, and residue variables). Variables that need to be initialized at the beginning of every iteration are all the others, which are related either to the construction of the layered residual networks (these are the reached, parent, in_previous_layer, and in_next_layer variables) or to the control of flow return (these are the stack, blocked, and returned variables). As a final observation on the variables employed by the algorithm, it should be noted that some of them are not used at all by some nodes, but do nonetheless appear listed for the sole sake of simplicity.

The reception of layer messages at \( n \) at a certain pulse in which reached = false causes parent, to point to one of the neighbors that sent the layer's. Each such neighbor \( n_i \) is sent an

\[
\text{ack} \text{ and in addition in}_{\text{previous}}_{\text{layer}}^i \text{ is set to true. Reception by n_i of an ack from n_i causes n_i to set in}_{\text{next}}_{\text{layer}}^i \text{ to true. Whenever a flow message is sent by n_i to n_i or received by n_i from n_i, the variables excess_i, flow_i, and residue_i are updated accordingly. When n_i cannot rid itself of its excess flow by sending it forward on the layered residual network, it returns that flow on a "last-in, first-out" basis (supported by stack) to the nodes that sent it. Nodes in the previous layer that do not get returned flow are sent a block message. Both returned flows and block messages signal the receiver that no more flow should during the current iteration be sent to n_i.}
\]

In Algorithm \text{S_Find_Max_Flow}, \( \mathcal{N}_k = \{n_i\} \) and \( 1 \leq k < K \). Again for the sake of simplicity (though at the expense of a longer algorithm), we have chosen to provide separate actions for \( n_i, n_i \in N - \{n_1, n_2\} \) and \( n_0 \).

\textbf{Algorithm S_Find_Max_Flow.}
Variables:
\[ \text{reached}_i, \text{parent}_i; \]
\[ \text{in}_{\text{previous layer}}_i \quad \text{for all } n_j \in \text{Neig}_i; \]
\[ \text{in}_{\text{next layer}}_i \quad \text{for all } n_j \in \text{Neig}_i; \]
\[ \text{stack}_i; \]
\[ \text{excess}_i = 0; \]
\[ \text{flow}_i^j = 0 \quad \text{for all } n_j \in \text{Neig}_i; \]
\[ \text{residue}_i^j = c(n_i, n_j) \quad \text{for all } (n_i \rightarrow n_j) \in E^n; \]
\[ \text{residue}_i^j = 0 \quad \text{for all } (n_j \rightarrow n_i) \in E^n; \]
\[ \text{blocked}_i^j = 0 \quad \text{for all } n_j \in \text{Neig}_i; \]
\[ \text{returned}_i^j = \text{for all } n_j \in \text{Neig}_i; \]
\[ y_i. \]

### Listing 7.10

**Input:**
\[ s = 4n(k - 1) \text{ or } s = 4n(K - 1), \text{ MSG}(s) = \emptyset \]

**Action if** \( n_i = n_1 \) (if \( n_i \in N_0 \), for \( k = 1 \)):
\[ \text{in}_{\text{next layer}}_i^j := \text{false} \quad \text{for all } n_j \in \text{Neig}_i; \]
\[ \text{blocked}_i^j := \text{false} \quad \text{for all } n_j \in \text{Neig}_i; \]
\[ K := k; \]
\[ \text{Send layer to all } n_j \in \text{Neig}_i \text{ such that } \text{residue}_i^j > 0. \]

### Listing 7.11

**Input:**
\[ s = 4n(k - 1) \text{ or } s = 4n(K - 1), \text{ MSG}(s) = \emptyset \]

**Action if** \( n_i \neq n_1 \neq n_0 \):
\[ \text{reached}_i := \text{true}; \]
\[ \text{parent}_i := \text{nil}; \]
in_previous_layer_i^j := false for all \( n_i \in \text{Neig}_i \);

in_next_layer_i^j := false for all \( n_i \in \text{Neig}_i \);

stack_i := \text{nil};

\text{blocked}_i^j := false for all \( n_i \in \text{Neig}_i \);

\text{returned}_i^j := false for all \( n_i \in \text{Neig}_i \);

\( K := k. \)

---

**Listing 7.12**

**Input:**

\[ s = 4n(k - 1) \text{ or } s = 4n(K - 1), \quad \text{MSG}(s) = \emptyset \]

**Action if** \( n_i = n_n \):

\[
\text{reached}_i := \text{true}; \\
\text{parent}_i := \text{nil}; \\
K := k. 
\]

---

**Listing 7.13**

**Input:**

\[ 4n(k - 1) + 1 \leq s \leq 4nk - 2n - 1 \text{ or } 4n(K - 1) + 1 \leq s \leq 4nK - 2n - 1, \quad \text{MSG}(s) \text{ such that } \text{origin}(msg) = (n_i, n_j) \text{ for } msg \in \text{MSG}(s). \]

**Action if** \( n_i = n_n \):

\[
\text{for all } \text{ack} \in \text{MSG}(s) \text{ do } \\
\text{in_next_layer}_i^j := \text{true}; \\
\text{if there exists } \text{success} \in \text{MSG}(s) \text{ then } \\
K := \infty. 
\]

---

**Listing 7.14**

**Input:**

\[ 4n(k - 1) + 1 \leq s \leq 4nk - 2n - 1 \text{ or } 4n(K - 1) + 1 \leq s \leq 4nK - 2n - 1, \quad \text{MSG}(s) \text{ such that } \text{origin}(msg) = (n_i, n_j) \text{ for } msg \in \text{MSG}(s). \]

**Action if** \( n_i \neq n_n \text{ and } n_i \neq n_n \):

\[
\text{if not } \text{reached}, \text{ then } \\
\text{if there exists } \text{layer} \in \text{MSG}(s) \text{ then } \\
\text{begin } \\
\text{reached}_i := \text{true}; \\
\text{end. } 
\]
for all layer ∈ $MSG(s)$ do
begin
if parent, = nil then
parent, := $n_1$;
\text{in}_\text{previous}_\text{layer}_i^j := \text{true};
Send ack to $n_j$
end;
Send layer to all $n_k ∈ \text{Neig}$ such that $n_k ≠ n_j$ and
\text{residue}_i^k > 0
end;
for all ack ∈ $MSG(s)$ do
\text{in}_\text{next}_\text{layer}_i^j := \text{true};
for success ∈ $MSG(s)$ do
begin
$K$ := $\infty$
Send success to parent,
end.

Listing 7.15

Input:
$4n(k - 1) + 1 ≤ s ≤ 4n_k - 2n - 1$ or $4n(k - 1) + 1 ≤ s ≤ 4nK - 2n - 1$, $MSG(s)$ such that origin$(msg) = (n_i, n_j)$ for $msg ∈ MSG(s)$.
Action if $n_i = n_n$:
if not reached, then
begin
if there exists layer ∈ $MSG(s)$ then
begin
reached, := \text{true};
for all layer ∈ $MSG(s)$ do
begin
if parent, = nil then
parent, := $n_j$;
Send ack to $n_j$
end
end
else
if $K = k$ then
begin
$K$ := $\infty$;
Send success to parent,
end.

Listing 7.16
Input:

\[ s = 4n_k - 2n, \quad \text{MSG}(0) = \emptyset \]

Action if \( n_i = n_n \):

\[
\text{for all } n_j \in \text{Neig}, \text{ such that } \quad \text{do} \]

\[
\begin{align*}
    y_i &:= \text{residue}_i^j; \\
    \text{flow}_i^j &:= \text{flow}_i^j + y_i; \\
    \text{residue}_i^j &:= 0; \\
    \text{Send } \text{flow}(y_i) \text{ to } n_i \\
\end{align*}
\]

end.

Listing 7.17

Input:

\[ 4n_k - 2n + 1 \leq s \leq 4n_k - 1, \quad \text{MSG}(s) \text{ such that } \text{origin}(msg) = (n_i, n_n) \text{ for } msg \in \text{MSG}(s). \]

Action if \( n_i = n_n \):

\[
\text{for all } \text{flow}(x) \in \text{MSG}(s) \text{ do} \]

\[
\begin{align*}
    \text{flow}_i^j &:= \text{flow}_i^j - x; \\
    \text{residue}_i^j &:= \text{residue}_i^j + x \\
\end{align*}
\]

end.

Listing 7.18

Input:

\[ 4n_k - 2n + 1 \leq s \leq 4n_k - 1, \quad \text{MSG}(s) \text{ such that } \text{origin}(msg) = (n_i, n_n) \text{ for } msg \in \text{MSG}(s). \]

Action if \( n_i \neq n_n \) and \( n_i \neq n_n \):

\[
\text{for all } \text{flow}(x) \in \text{MSG}(s) \text{ do} \]

\[
\begin{align*}
    \text{excess}_i &:= \text{excess}_i + x; \\
    \text{flow}_i^j &:= \text{flow}_i^j - x; \\
    \text{residue}_i^j &:= \text{residue}_i^j + x; \\
\end{align*}
\]

if \( \text{in}_{-\text{previous}}_{\text{layer}}^j_i \) then

Push \((n_i, x)\) onto stack;
if \( \text{in}_j \) \( \text{next}_j \) \( \text{layer}_j \) then

\( \text{blocked}_j := \text{true} \)

end;

for all \( \text{block} \in \text{MSG}(s) \) do

\( \text{blocked}_j := \text{true} \);

while (there exists \( n_k \in \text{Neig}_i \) such that

\( \text{blocked}_k \)) and \( \text{excess}_i > 0 \) do

begin \( y_i := \text{min} \{ \text{excess}_i, \} \);\n
\( \text{excess}_i := \text{excess}_i - y_i \);

\( \text{flow}_i := \text{flow}_i + y_i \);

\( \text{residue}_i := \text{residue}_i - y_i \);

Send \( \text{flow}(y) \) to \( n_k \)

end;

while \( \text{excess}_i > 0 \) do

begin Pop \( (n_k, x) \) off \( \text{stack}_i \);

\( y_i := \text{min} \{ \text{excess}_i, x \} \);

\( \text{excess}_i := \text{excess}_i - y_i \);

\( \text{flow}_i := \text{flow}_i + y_i \);

\( \text{residue}_i := \text{residue}_i - y_i \);

\( \text{returned}_i := \text{true} \);

Send \( \text{flow}(y) \) to \( n_k \)

end;

if there exists \( n_k \in \text{Neig}_i \) such that \( \text{returned}_i \) then

for all \( n_k \in \text{Neig}_i \) such that \( \text{returned}_i \) and not

\( \text{returned}_i \) do

begin \( \text{returned}_i := \text{true} \);

Send block to \( n_k \)

end.

Listing 7.19

Input:
4n_s - 2n + 1 \leq s \leq 4n_s - 1, MSG(s) such that \text{origin}_s (msg) = (n_s, n_i) for msg \in MSG(s).

**Action if** \text{n}_i = \text{n}_s:

**for all** \text{flow}(x) \in MSG(s)**

**begin**

\[
\begin{align*}
\text{flow}_i^j & := \text{flow}_i^j - x, \\
\text{residue}_i^j & := \text{residue}_i^j + x
\end{align*}
\]

**end.**

In Algorithm S\_Find\_Max\_Flow, (7.10) through (7.19) realize the \text{K}\_iteration, two-phase-per-iteration method that we described. Actions (7.10) through (7.15) handle the first phases of the \text{K} iterations, while actions (7.16) through (7.19) handle the second phases of the \text{K} - 1 first iterations. Actions (7.10) through (7.12) last for exactly one pulse per iteration each, and are intended respectively for \text{n}_i, \text{n}_j \in \text{N} - \{\text{n}_s, \text{n}_k\}, and \text{n}_s to initialize their variables for the new iteration and for \text{n}_i to send out the initial layer messages. Actions (7.13) through (7.15) are executed for 2n - 1 pulses each in every iteration, and specify the participation of the nodes in the remainder of the first phase. What these actions contain are the responses, respectively by \text{n}_i, \text{n}_j, and \text{n}_s, to the receipt of layer, ack, and success messages. Note, however, that \text{n}_i never receives a layer and \text{n}_s never receives an ack or success. It is through (7.15) that a success message first gets sent in each iteration.

In (7.16), which is executed for exactly one pulse in each iteration, \text{n}_i sends out the initial flow messages of the iteration. The handling of such messages, and of block messages, is achieved through (7.17) through (7.19) for nodes \text{n}_i, \text{n}_j \in \text{N} - \{\text{n}_s, \text{n}_k\}, and \text{n}_s, respectively (\text{n}_s never receives any block message, though). Each of these actions lasts for 2n - 1 pulses, and is tuned to the peculiarities of the corresponding node or nodes. Specifically, (7.17) and (7.19), for execution respectively by \text{n}_j and \text{n}_s, do not include the sending of any messages at all. Also, in (7.17) \text{n}_i does not act upon the receipt of block messages. Note that no action is explicitly given for nodes \text{n}_i \neq \text{n}_j, at pulse \text{s} = 4\text{n}_s - 2\text{n}, as this is the first pulse in the second phase of iteration \text{k} and is as such meant for \text{n}_i only.

When a new iteration is initiated in (7.10) through (7.12) and the pertinent variables get initialized, nodes also set \text{K} to the number \text{k} of the current iteration, in preparation for the possibility that this may be the last one. When \text{n}_s sends a success message in (7.15), or upon receipt of such a message by \text{n}_s or \text{m}_s \in \text{N} - \{\text{n}_s, \text{n}_s\} respectively in (7.13) and (7.14), \text{K} is reset to infinity, thereby indicating that the current iteration will include a second phase, and then is not the last one. When the algorithm terminates, the flow variables contain a maximum flow in \text{G}_\text{s}.

The other synchronous algorithm that we study in this section is considerably simpler than Algorithm S\_Find\_Max\_Flow. For this reason, we only describe it superficially, and leave to the reader the task of expressing it more formally in the style of notation we have been employing (cf. Exercise 4). This algorithm works with the notion of preflows, and continually tries to push excess flow along the edges of the residual network that the nodes estimate to be on the shortest paths to \text{n}_i or \text{n}_j. The algorithm starts with a preflow \text{f} such that \text{f}(\text{n}_s, \text{n}_i) = \text{c} (\text{n}_s, \text{n}_i) for all (\text{n}_s \rightarrow \text{n}_i) \in \text{E}_s and \text{f}(\text{n}_i, \text{n}_s) = 0 for all (\text{n}_i \rightarrow \text{n}_s) \in \text{E}_s, with \text{n}_s \neq \text{n}_i.

Every node \text{n}_i maintains an estimate \text{d}_i of its shortest distance to either \text{n}_i or \text{n}_i, in the residual network. Any initial values for these estimates will do, as long as \text{d}_i = \text{n}_i, \text{d}_i = \text{0}, and \text{d}_i \leq \text{d}_i + 1 for all (\text{n}_j \rightarrow \text{n}_i) \in \text{E}_s. In the first pulse of the algorithm, these estimates are exchanged between neighbors; at node \text{n}_i, the estimate of neighbor \text{n}_j is stored in \text{d}_j^i. At all times during the execution of the algorithm, these estimates are such that either \text{d}_i is a lower bound on the distance from \text{n}_i to \text{n}_i, if \text{d}_i < \text{n}_i, or \text{d}_i - \text{n}_i is a lower bound on the distance from \text{n}_i to \text{n}_i, if \text{d}_i \geq \text{n}_i.
At each pulse of the algorithm, the active nodes attempt to get rid of their excess flows by pushing flow in the direction of _n_i_ or _n_j_. Letting _f_ be the preflow at the end of the previous pulse (the initial preflow, in the first pulse), a node _n_i_ is said to be active if _n_i_ /∈ _N_ \ {_m_i_, _n_j_} and _ef(_n_i_) > 0. An active node _n_i_ at the current pulse, first sends an amount of flow equal to 

\[
\min_{j \in \{m_i, n_j\}} d_i^j + 1
\]

to a neighbor _n_j_ such that _d_i^j = d_i^j + 1_ and _c_f(_n_i_, _n_j_) > 0_, and updates _f_ (as well as _ef(_n_i_) and _cf(_n_i_, _n_j_)_ accordingly. This is repeated until either _e_f(_n_j_, _n_i_) = 0_ or _c_f(_n_j_, _n_i_) = 0_ for all _n_j_ such that _d_i^j = d_i^j + 1_. If after this _e_f(_n_j_, _n_i_) > 0_, then _d_i_ is updated to the minimum, over all neighbors _n_j_ of _n_i_ such that _c_f(_n_j_, _n_i_) > 0_, of _d_i^j_ + 1_, and this value, if different from the previous one, is sent to _n_j_’s. The next pulse is initiated by adding to _e_f(_n_i_) all the flow received during the pulse. The algorithm terminates when no nodes are any longer active, although a termination criterion that, as in previous occasions, only considers the number of pulses elapsed is also possible.

7.2.3 An asynchronous algorithm

In this section, we discuss briefly the asynchronous version of the preflow-based synchronous algorithm that we introduced in the previous section. As with its synchronous counterpart, we leave all the details for the reader to pursue as an exercise (cf. Exercise 5).

The essential difficulty in the asynchronous case is that the condition that _d_i = d_i^j + 1_ is necessary for _n_i_ to send flow to _n_j_ (cf. Section 7.2.2 on the preflow-based synchronous algorithm), cannot be trivially ensured, as the values of _d_i_ and _d_j^i_ may differ substantially. The solution adopted when proposing the corresponding asynchronous algorithm has been that every flow sent from _n_i_ to _n_j_ must carry the value of _d_i_ , and be explicitly accepted or rejected by _n_j_ before additional flow may be sent. When _n_j_ receives flow from _n_i_ and verifies that in fact _d_i = d_i^j + 1_, then the flow is accepted and this is reported back to _n_i_. If, on the other hand, _d_i ≠ d_i^j + 1_, then the flow is rejected and this is reported back to _n_i_ along with the value of _d_i_. Upon receiving this rejection message, _n_i_ updates _e_f(_n_i_), _c_f(_n_i_, _n_j_), _d_i^j_, and possibly _d_i_. Whenever _d_i_ changes, its new value is reported to all of _n_i_’s neighbors.

7.3 Exercises

1. Prove properties (i) and (ii) of Section 7.1.1 on minimum spanning trees.
2. Discuss how to modify Algorithm _A_Find_MST_ for the case in which edges are not FIFO. In particular, show that the situations in which a connect or test message cannot be replied to immediately can be handled with the aid of conditions for actions to be executed, instead of auxiliary variables.
3. In the context of Section 7.2.1, show that the definitions of the value of _f_ as the total flow going out from _n_1_ or coming into _n_n_ are indeed equivalent to each other.
4. Express the second synchronous algorithm of Section 7.2.2 according to Algorithm _S_Template_.
5. Express the asynchronous algorithm of Section 7.2.3 according to Algorithm _A_Template_.

1. Prove properties (i) and (ii) of Section 7.1.1 on minimum spanning trees.
2. Discuss how to modify Algorithm _A_Find_MST_ for the case in which edges are not FIFO. In particular, show that the situations in which a connect or test message cannot be replied to immediately can be handled with the aid of conditions for actions to be executed, instead of auxiliary variables.
3. In the context of Section 7.2.1, show that the definitions of the value of _f_ as the total flow "going out from" _n_1_ or "coming into" _n_n_ are indeed equivalent to each other.
4. Express the second synchronous algorithm of Section 7.2.2 according to Algorithm _S_Template_.
Express the asynchronous algorithm of Section 7.2.3 according to Algorithm A_Template.

### 7.4 Bibliographic notes

Our treatment in Section 7.1 of the problem of finding a minimum spanning tree follows the original paper of Gallager, Humblet, and Spira (1983) closely, except for the material in Section 7.1.3, which is based on Gafni (1985) and Chin and Ting (1990). For an algorithm with time complexity even lower than the one mentioned in Section 7.1.3, the reader is referred to Awerbuch (1987). Another publication of interest is Janssen and Zwiers (1992). For material on maximum flows in networks to complement our treatment in Section 7.2.1, the reader can count on books dedicated exclusively to the subject (Ford and Fulkerson, 1962; Ahuja, Magnanti, and Orlin, 1993), chapters in more general books (Lawler, 1976; Even, 1979; Papadimitriou and Steiglitz, 1982; Cormen, Leiserson, and Rivest, 1990), and surveys (Ahuja, Magnanti, and Orlin, 1989; Goldberg, Tardos, and Tarjan, 1990).

Algorithm S_Find_Max_Flow of Section 7.2.2 is from Awerbuch (1985b), and the concepts of augmenting paths and of layered residual networks that it employs are originally from Ford and Fulkerson (1962) and Dinic (1970), respectively. The algorithm in Awerbuch (1985b) is an adaptation of the algorithm given by Shiloach and Vishkin (1982) for a shared-memory model (Karp and Ramachandran, 1990). The other synchronous algorithm of Section 7.2.2 and the asynchronous algorithm of Section 7.2.3 can be found in detail in Goldberg and Tarjan (1988). The concept of preflows on which they are based is originally from Karzanov (1974).

Parallel implementations of the algorithms of Goldberg and Tarjan (1988) have been discussed by Anderson and Setubal (1992) and by Portella and Barbosa (1992). In the latter publication, the authors describe an experimental evaluation of all the three algorithms discussed in Section 7.2. This evaluation employs random graphs (Bollobás, 1985) in the style suggested in DIMACS (1990).

### Chapter 8: Resource Sharing

#### Overview

When the nodes in $G$ share resources with one another that must not be accessed by more than one node at the same time, distributed algorithms must be devised to ensure the mutual exclusion in the access to those resources, that is, ensure that nodes exclude one another in time to access the shared resources. This problem is not entirely new to us, having been treated in Section 6.1 in the context of self-stabilization on a ring, and in Section 6.3 in the context of detecting deadlocks in a distributed computation in which nodes provide service to one another, but never to more than one node at a time.

In this chapter, $G$ is an undirected graph, and our treatment spans two main problems. The first problem is to ensure mutual exclusion when all the nodes share one single resource, or a group of resources that always have to be accessed as a single one. In this case, $G$ may be as dense as a complete graph, reflecting the need, in some algorithms, for a node to communicate with all others to secure exclusive resource access. Mentions in the literature to the "mutual exclusion problem" normally refer to this first problem, which we address in Section 8.1.

The second problem that we treat in this chapter is that of ensuring mutual exclusion when each node may require access to a different set of resources. When a node accesses the same set of resources whenever it accesses any resource, the problem is a generalized form of the paradigmatic dining philosophers problem. When the set of resources that a node accesses may vary from one time to the next, then the problem has become known as the drinking philosophers problem. We dedicate Sections 8.3 and 8.4 respectively to each of these problems, after a common introduction in Section 8.2.

Two important notions that pervade all of our resource sharing studies in this chapter are those of a deadlock, which already we are acquainted with, and of starvation. Acceptable algorithms for resource sharing must ensure that neither conditions are ever present, unless it can be argued, in the particular situation at hand, that resorting to deadlock detection is preferable, as we discussed in Section 6.3. In the context of ensuring mutual exclusion in the access to shared resources by the group $N$ of nodes, deadlock exists when none of the
nodes ever succeeds in obtaining access to the resources. If there always exists at least one node that does succeed, but at least one other node does not succeed indefinitely, then the situation is one of starvation.

Exercises and bibliographic notes are given, respectively, in Sections 8.5 and 8.6.

8.1 Algorithms for mutual exclusion

In this section, as in other occasions in the book, we assume that nodes have distinct identifications totally ordered by <. For node $n$, such an identification is $id$. Nodes share a resource, or a group of resources, that must be accessed with the guarantee of mutual exclusion. If it is a group of resources that the nodes share, then we assume that all the resources in the group are always accessed together, as if they constituted one single resource, so that for all purposes it is legitimate to assume that the nodes share one single resource.

For the first algorithm that we study, $G$ is a complete graph, because the algorithm is based on the strategy that a node, in order to access the shared resource, must obtain permission to do so from all the other $n-1$ nodes. This first algorithm is called $A_{Mutually\_Exclude\_C}$ (the suffix "C" here indicates, as in Section 5.1, that a complete graph is involved), and is based on the following simple approach. In order to request permission to access the shared resource, node $n$ sends a $request(seq, id)$ message to all the other nodes in $G$. The parameters that this message carries are, respectively, a "sequence number" (akin to the tag attached to marker's in Algorithm $A_{Detect\_Termination}$ of Section 6.2.1) and $n$'s identification. The sequence number is an integer, and is obtained by adding one to the largest such number $n$ has received or sent in a request message (or to zero, if no request has ever been received or sent by it). Node $n$ proceeds to access the resource upon receiving one reply message from each of the other nodes.

Upon receiving a request $(seq, id)$, node $n$ replies immediately to $n$, with a reply message if it is not waiting for reply's itself. If it is waiting for reply's then it is also competing for exclusive access to the shared resource, and the parameters that it sent out with its request messages, namely a sequence number $seq'$ and $id$, must be compared to those received with the message from $n$ to determine which node takes priority. Lower sequence numbers indicate earlier request messages (in the sense of the partial order $≺$ of Section 3.1), so that $n_j$ takes priority (i.e., is sent a reply by $n$) if

$$seq_j = seq_i \land \left( id_j = id_i \lor \right.$$}

where the comparison is done lexicographically. Otherwise, $n$ delays the sending of a reply to $n_j$, until after it has accessed the shared resource.

In Algorithm $A_{Mutually\_Exclude\_C}$, the following are the variables employed by node $n$.

Two integers, $seq$ and $highest\_seq$ (the latter initialized to zero), are used respectively to indicate the sequence number $n$ sent with the last group of request messages it sent (if any) and the highest sequence number to have been sent or received by $n$ in a request message. Another integer, $expected$, (set to zero initially), indicates the number of reply messages $n$ must receive before accessing the shared resource. For all $n_j \in Neig$, a Boolean

$owes\_reply_j^i = false$ (initially set to $false$) is used to indicate whether $n_i$ has postponed the sending of a reply to $n_j$.

Algorithm $A_{Mutually\_Exclude\_C}$:

Variables:

- seq;
- highest_seq = 0;
- expected = 0;

$owes\_reply_j^i = false$ for all $n_j \in Neig$. 

Listing 8.1
Input:
   \( msg_i = \text{nil}. \)
Action when \( \text{expected}_i = 0 \) and access to the shared resource is needed:
   \[
   \begin{align*}
   \text{seq}_i & := \text{highest}_\text{seq}_i + 1; \\
   \text{highest}_\text{seq}_i & := \text{seq}_i; \\
   \text{expected}_i & := n - 1; \\
   \text{Send request}(\text{seq}_i, \text{id}_i) \text{ to all } n_j \in \text{Neig}. 
   \end{align*}
   \]

Listing 8.2

Input:
   \( msg_i = \text{request}(\text{seq}, \text{id}_i) \) such that \( \text{origin}(msg_i) = (n_i, n_j). \)
Action:
   \[
   \begin{align*}
   \text{highest}_\text{seq}_i & := \max\{\text{highest}_\text{seq}_i, \text{seq}\}; \\
   \text{if } \text{expected}_i & = 0 \text{ or } (\text{seq}, \text{id}_i) < (\text{seq}_i, \text{id}_i) \text{ then} \\
   & \text{Send reply to } n_j \\
   \text{else} \\
   \text{owe_reply}_i^j & := \text{true}. 
   \end{align*}
   \]

Listing 8.3

Input:
   \( msg_i = \text{reply}. \)
Action:
   \[
   \begin{align*}
   \text{expected}_i & := \text{expected}_i - 1; \\
   \text{if } \text{expected}_i & = 0 \text{ then} \\
   & \text{begin} \\
   & \text{Access shared resource;} \\
   & \text{for all } n_j \in \text{Neig} \text{ such that} \\
   & \text{begin} \\
   & \text{owe_reply}_i^j := \text{false}; \\
   & \text{Send reply to } n_j \\
   & \text{end} \\
   & \text{end} \\
   \end{align*}
   \]

In Algorithm \( A\_Mutually\_Exclude\_C \), actions (8.1) through (8.3) indicate \( n_i \)'s participation respectively as a spontaneous initiator, upon receipt of a \( \text{request} \), and upon receipt of a \( \text{reply} \). Action (8.1), in particular, is executed whenever there is need for \( n_i \) to access the shared resource and in addition \( \text{expected}_i = 0 \) (indicating that \( n_i \) is not already engaged in seeking
access to that resource). As in Algorithm $A\_\text{FIFO}$ of Section 2.1, a node may initiate its participation in the algorithm spontaneously more than once, in that case by deciding to migrate a task to run elsewhere, in this case by deciding that access to the shared resource is needed. What this amounts to is that $N_0$ must be regarded as a maximal set of nodes that send out request's concurrently. Each such set initiates a new execution of the algorithm, and executions operate on variables that persist, in the sense of not being re-initialized, from one execution to another. Theorem 8.1 establishes important properties of this algorithm.

**Theorem 8.1.**
Algorithm $A\text{\_Mutually\_Exclude\_C}$ ensures mutual exclusion in the access to the shared resource, and is in addition deadlock-and starvation-free.

**Proof:** Two nodes can only access the shared resource concurrently if they receive the $n - 1$st reply message concurrently. This follows from (8.3) and, in particular, indicates that each of the two nodes must have received a reply from the other as well. But by (8.2) and (8.3), and because node identifications are all distinct from one another, this can only have happened if at least one of the two was not requesting access to the resource, which is in contradiction with the possibility that they access the resource concurrently.

By (8.2), node $n_i$ only refrains from sending $n_j$ a reply if expected $> 0$ and $(\text{seq}_i, \text{id}_i) \prec (\text{seq}_j, \text{id}_j)$, where seq and id are the parameters of $n_i$'s request message. In this case, $n_j$ is forced to wait for $n_i$'s reply. Because node identifications are totally ordered by $<$, a wait cycle cannot be formed among the nodes, and then no deadlock can ever occur (cf. Section 6.3). Now consider the number of resource accesses that may take place after node $n_i$ has sent request's and before it has received reply's (because mutual exclusion is ensured, resource accesses are totally ordered, so that the "before" and "after" are meaningful with respect to this order). By (8.1) and (8.2), the sequence number a node sends along with a request message is strictly greater than those it has received in request's itself, so that by (8.2) every node sending out request's after receiving $n_i$'s request will only access the shared resource after $n_i$ has done so. The number of resource accesses we are considering is then finite, and as a consequence no starvation ever occurs.

Let us now examine the complexities of Algorithm $A\text{\_Mutually\_Exclude\_C}$. Clearly, each access to the shared resource involves $n - 1$ request messages and $n - 1$ reply's. The algorithm's message complexity per access to the shared resource is then $O(n)$. The time complexity per access to the shared resource refers to the chain of messages that may occur starting with a request sent by a node and the last reply that it receives. The longest such chain occurs when a global state exists in which $n - 1$ nodes in a row have withheld reply's from the next node in the sequence (by Theorem 8.1, the number of nodes involved in this wait cannot be greater than $n - 1$, because otherwise there would be deadlock). If $n_i$ and $n_j$ are, respectively, the first and last nodes in this wait chain (that is, $n_i$ is the only node not to have withheld a reply from), and if the request from $n_i$ arrives at $n_j$ before $n_j$ accesses the shared resource, then the reply's that $n_j$ sends out when it finally does access the resource start a causal chain of reply's through the other nodes to $n_i$. The time complexity of the algorithm per access to the shared resource is then $O(n)$ as well.

The algorithm's bit complexity is in principle unbounded, because, although a node's identification can be as usual assumed to be expressible in $\log n$ bits, the other parameter that request messages carry, the sequence number, does not in the algorithm have any bound. However, it can be argued relatively easily that no two sequence numbers that a node has to compare in (8.2) are ever farther apart from each other than $n - 1$, and then it is possible to implement them as $O(\log n)$-bit numbers (cf. Exercise 1). The algorithm's bit complexity is then $O(n \log n)$.

Unlike Algorithm $A\text{\_Mutually\_Exclude\_C}$, the next algorithm that we consider in this section does not require every node to receive explicit permission from every other node before accessing the shared resource. What makes this possible is the following interesting observation. For $n \in N$, let $S_i \subseteq N$ denote the set of nodes from which $n_i$ must receive explicit permission before accessing the shared resource (in the previous algorithm, $S_i = N \setminus \{n_i\}$ for all $n_i \in N$). In order for $S_i$ not to have to include every node in $N \setminus \{n_i\}$ yet mutual exclusion to be ensured in the access to the shared resource when the nodes in $S_i$ grant
permission for $n$ to proceed, for every two nodes $n$ and $n'$ we must have $S_i \cap S_i' \neq \emptyset$. If the $S$ sets can be built such that this property holds, then every pair of conflicting requests to access the shared resource will reach at least one node, which will then be able to arbitrate between the two requests and ensure mutual exclusion.

Once the sets $S_1,...,S_n$ have been determined, the following is how a node $n$ proceeds in order to access the shared resource. For simplicity when describing the algorithm, we assume that $n \notin S_i$, so that the number of request's that $n$ sends is $|S_i|$. First $n$ sends a request message to every node in $S_i$ and then waits to receive one granted message corresponding to each of the request messages it sent. It may then access the shared resource, and after doing so sends a release message to each of the nodes in $S_i$. A node that has sent a granted and receives another request before receiving the corresponding release's must somehow postpone its permission corresponding to the new request. Although the granted messages may be thought of as corresponding to the reply messages of the previous algorithm, the need to explicitly indicate that the resource is no longer in use through the release messages reflects some of the important differences between the two approaches. The essential reason why release messages are now needed is that a request does not reach every node, and thence the double meaning that a reply message had of both granting permission and signaling the end of an access to the shared resource can no longer be exploited with the granted messages. In fact, another consequence of the selective broadcast of request's in addition to the need of explicit release's is that deadlocks can no longer be taken as prevented even if the request's carry the same information that they did in the previous case. Because different nodes obtain their permissions from different sets of nodes, it is rather simple to imagine situations in which wait cycles appear.

The following is then the overall strategy to handle conflicts and the waits that result from them, and yet ensure that deadlocks are not possible. A request message is, as in the previous algorithm, sent by $n$ as request(seq, id), where seq is strictly greater than every other sequence number $n$ has ever sent or received in such a message. Node $n$ maintains a Boolean variable, called locked, and initialized to false, to indicate whether it has sent a granted message to some node without having received any other message from that node. When $n$ receives a request(seq, id) and locked = false, a granted is immediately sent to its originator. If locked = true, then $n$ marks the origin of the request message for later consideration. Upon delaying the response to a node in this way, $n$ must ensure that no deadlock will arise, and to this end proceeds as follows. If the newly received request takes precedence (in the sense of a lexicographically smaller pair (seq, id)) over the request to which $n$ has replied with a granted as well as all the others that $n$ has not yet replied to, then a probe message is sent to the same node to which the granted was sent. Otherwise, a delayed message is sent in response to the new request, which is then kept waiting. A node that receives a probe responds to it right away with a relinquish if it has already received a delayed, or when it does receive a delayed, if it still has not. Node $n$ does not send another probe until a relinquish or a release has arrived for the one it has sent. A node only sends a relinquish in response to a probe if a granted was not received from each of the nodes that sent it a delayed.

Algorithm A_Mutually_Exclude, presented next, is based on this approach. In contrast with the previous approach, $G$ is no longer a complete graph, but rather has its set of edges given in accordance with the sets $S_1,...,S_n$ in such a way that $(n,n') \in E$ if and only if $n, n' \in S_i$ or $n, n' \in S_j$. Also, we assume that all edges are FIFO, and then a granted never overruns a delayed or a probe a granted.

In addition to the variables seq, highest_seq, and expected, used here as in the previous algorithm, and the already introduced variable locked, Algorithm A_Mutually_Exclude employs the following additional variables. A request in response to which a granted has been sent has its origin and seq and id parameters recorded by $n$ in the variables granted_node, granted_seq, and granted_id, respectively. Node $n$ maintains a queue, called queue, and initialized to nil, to store these same attributes for all request's that cannot be immediately replied to. This queue is maintained in lexicographically increasing order of (seq, id). Finally, the Booleans has_probed, and got_probe are
for all \( n_j \in S_i \), all initialized to false, are employed to indicate respectively whether \( n_i \) has sent a probe for which a relinquish or a release was not received in response, whether \( n_i \) has received a probe from \( n_j \), and whether a delayed was received from \( n_j \) without a succeeding granted.

**Algorithm A. Mutually Exclude:**

Variables:

\[
\begin{align*}
\text{seq}_i & = 0; \\
\text{highest_seq}_i & = 0; \\
\text{expected}_i & = 0; \\
\text{locked}_i & = \text{false}; \\
\text{granted_node}_i & ; \\
\text{granted_seq}_i & ; \\
\text{granted_id}_i & ; \\
\text{queue}_i & = \text{nil}; \\
\text{has_probed}_i & = \text{false}; \\
\text{gotprobe}_i & = \text{false} \text{ for all } n_j \in S_i ; \\
\text{gotdelayed}_i & = \text{false} \text{ for all } n_j \in S_i .
\end{align*}
\]

**Listing 8.4**

Input:

\( msg_i = \text{nil} \).

Action when \( \text{expected}_i = 0 \) and access to the shared resource is needed:

\[
\begin{align*}
\text{seq}_i & := \text{highest_seq}_i + 1; \\
\text{highest_seq}_i & := \text{seq}_i; \\
\text{expected}_i & := |S_i|; \\
\text{Send request(seq}_i, \text{id}_i) \text{ to all } n_j \in S_i .
\end{align*}
\]

**Listing 8.5**

Input:

\( msg_i = \text{request(seq, id)} \) such that \( \text{origin}(msg_i) = (n_i, n_j) \).

Action:

\[
\begin{align*}
\text{highest_seq}_i & := \text{max}(|\text{highest_seq}_i, \text{seq}|); \\
\text{if not locked, then} \\
\\text{begin} \\
\\text{locked}_i & := \text{true}; \\
\\text{granted_node}_i & := n_j; \\
\\text{granted_seq}_i & := \text{seq}; \\
\\text{granted_id}_i & := \text{id}; \\
\\text{Send granted to } n_j \\
\\text{end} \\
\text{else}
\end{align*}
\]
begin
Add \((n_j, \text{seq}, \text{id})\) to queue;  
if \((\text{seq}, \text{id}) < (\text{granted_seq}_i, \text{granted_id}_i)\) and \((n_j, \text{seq}, \text{id})\) is  
first in queue, then  
begin  
if not has_probed, then  
begin  
\text{has_probed} := \text{true};  
Send probe to \text{granted_node}_i;  
end  
end  
else  
Send delayed to \(n_j\)  
end.

Listing 8.6

Input:  
msg\(_i\) = \text{granted} such that \text{origin}(\text{msg}) = (n_i, n_j).  
Action:  
\text{expected} := \text{expected} - 1;  
if \(\text{got_delayed}_i^j\) then  
\(\text{got_delayed}_i^j := \text{false};\)  
if \text{expected} = 0 then  
begin  
Access shared resource;  
for all \(n_k \in \text{S}\) such that \(\text{got_probe}_i^k\) do  
\(\text{got_probe}_i^k := \text{false};\)  
Send release to all \(n_k \in \text{S}\);  
end.

Listing 8.7

Input:  
\text{msg} = \text{release}.  
Action:  
if \text{has_probed} then  
\text{has_probed} := \text{false};  
if queue\(_i\) = \text{nil} then  
\text{locked} := \text{false}  
else  
begin  
Let \((\text{granted_node}_i, \text{granted_seq}_i, \text{granted_id}_i)\) be first in)
queue;
Remove (granted_node, granted_seq, granted_id) from queue;
Send granted to granted_node;
end.

Listing 8.8

Input:
\[ msg_i = \text{probe such that origin}(msg_i) = (n_i, n_j). \]

Action:
\[
\text{if there exists } n_k \in S_i \text{ such that } \exists \text{ got}_{-}\text{delayed}_i^k \text{ then begin }
\quad \text{expected}_i := \text{expected}_i + 1;
\quad \text{Send relinquish to } n_j
\quad \text{end else } \quad \text{got}_{-}\text{probe}_i^j := \text{true.}
\]

Listing 8.9

Input:
\[ msg_i = \text{delayed such that origin}(msg_i) = (n_i, n_j). \]

Action:
\[
\quad \text{got}_{-}\text{delayed}_i^j := \text{true};
\quad \text{for all } n_k \in S_i \text{ such that } \exists \text{ got}_{-}\text{probe}_i^k \text{ do begin }
\quad \quad \text{expected}_i := \text{expected}_i + 1;
\quad \quad \text{Send relinquish to } n_k
\quad \quad \text{end.}
\]

Listing 8.10

Input:
\[ msg_i = \text{relinquish.} \]

Action:
\[
\quad \text{has_probed}_i := \text{false};
\quad \text{Add (granted_node, granted_seq, granted_id) to queue;}
\quad \text{Let (granted_node, granted_seq, granted_id) be first in queue;}
\quad \text{Remove (granted_node, granted_seq, granted_id) from queue;}
\]

Actions (8.4) through (8.10) realize the algorithm we described informally earlier. In this algorithm, the set $N_i$ is to be interpreted as in the case of Algorithm $A_{Mutually\_Exclude\_C}$. Some of the algorithm's properties are established by the following theorem.

**Theorem 8.2.**
Algorithm $A_{Mutually\_Exclude}$ ensures mutual exclusion in the access to the shared resource, and is in addition deadlock- and starvation-free.

**Proof:** By (8.4) and (8.6), any two nodes $n_i$ and $n_j$ can only access the shared resource concurrently if they receive, respectively, the $|S_i|$th and $|S_j|$th granted messages concurrently. However, by definition $S_i$ and $S_j$ have at least one node in common, say $n_k$, which by (8.5), (8.7), and (8.10) only sends granted messages when locked$_i$ = false upon receipt of a request or when a release or a relinquish is received. In addition, locked$_i$ is only false initially or upon receipt by $n_k$ of a release, so that either $n_k$ receives a release or a relinquish from $n_i$ before sending a granted to $n_i$, or conversely. In either case, a contradiction exists with the possibility of concurrent access to the shared resource by $n_i$ and $n_j$.

Different nodes send request's to different subsets of $N$, a wait cycle (cf. Section 6.3) may indeed be formed, but only momentarily though, because (8.5) ensures that a request(seq, id) arriving at $n_i$ prompts the sending by $n_i$ of a probe if (seq, id) is lexicographically minimum among (granted_seq, granted_id) and all the pairs in queue. All node identifications are totally ordered by <, and for this reason at least one probe must succeed in breaking the wait cycle through the sending of a relinquish by (8.8) or (8.9). Deadlocks are then not possible.

Ensuring mutual exclusion has the effect that all accesses to the shared resource are totally ordered. With respect to this total order, it is then legitimate to consider the number of accesses that may take place after an access by node $n_i$ and before the next access by the same node. Considering that in all nodes queues are kept in increasing lexicographic order of sequence numbers and node identifications, and that nodes issue request's with strictly increasing sequence numbers, the number of accesses we are considering is then clearly finite, and therefore starvation never occurs.

Analyzing Algorithm $A_{Mutually\_Exclude}$ for its complexity measures requires that we consider the sets $S_1, \ldots, S_n$ more closely, because, by the algorithm's actions, the message complexity per access to the shared resource by node $n_i$ is intimately dependent upon the number of nodes in $S_i$. Before proceeding any further with this discussion, though, we make two additional assumptions concerning these sets. The first assumption is that $|S_i| = |S_j|$ and the second assumption is that every node is contained in the same number of sets. The first assumption seeks to guarantee that every node requests an equal number of permissions in order to access the shared resource, while the second assumption aims at distributing evenly the "responsibility" for granting permission to access the shared resource.

Combined, the two assumptions are to be regarded as an attempt at "fairness," however this may be defined.

If we let $K$ denote the size of each of the $n$ sets and $D$ the number of sets of which each node is a member, then we have

$$n = \frac{|S_1| + \cdots + |S_n|}{D},$$

which yields $D = K$.

One of the possibilities for the sets $S_1, \ldots, S_n$ is of course $S_i = N$ for all $n_i \in N$. It is a trivial matter, in this case, to simplify Algorithm $A_{Mutually\_Exclude}$ until Algorithm $A_{Mutually\_Exclude\_C}$ is obtained, as $G$ is clearly a complete graph. Our interest, however, is in obtaining the smallest sets whose pairwise intersections are nonempty, because this is what will improve the algorithm's message complexity from that of Algorithm $A_{Mutually\_Exclude\_C}$.
Now consider the set $S_i$ for some node $n_i$. This set has $K$ members, each of which belonging to other $D - 1$ sets, so the number of distinct sets involving the members of $S_i$ is at most $K(D - 1) + 1$. Because we need $n$ such sets ($S_1$ through $S_n$), and considering that $D = K$ and that the largest number of sets leads to the smallest sets, we then have

$$n = K(K - 1) + 1,$$

which immediately yields $K \geq \sqrt{n}$. If $n$ is the square of some integer, then we adopt $|S_1| = \cdots = |S_n| = \sqrt{n}$.

Otherwise, a little imbalance cannot be avoided as some sets must necessarily have more nodes than the others, although no more than the square root of the least perfect square greater than $n$. It should be clear that this square root is still $O(\sqrt{n})$ lower bound for $K$, the message complexity of Algorithm $A_{Mutually\_Exclude}$ can be seen to be of $O(\sqrt{n})$ per access to the shared resource. This is so because, in the worst case, there will be $\sqrt{n}$ request messages, each one causing a probe message, this one being replied to by a relinquish message, in turn generating a granted message and eventually a release message. Notice that the argument we employed earlier on the possibility of bounding sequence numbers in the case of Algorithm $A_{Mutually\_Exclude\_C}$ is no longer applicable, so that the bit complexity is in this case unbounded (cf. Exercise 2).

As in the case of the previous algorithm, the time complexity of Algorithm $A_{Mutually\_Exclude}$ per access to the shared resource refers to the chain of messages that may occur starting with a request sent by a node and the release's that it sends. One possibility for this chain occurs in the following situation. Node $n_i$ sends $n$, a request, which is queued and only replied to with a granted after $n_i$ has sent a granted and received a release for each of the requests ahead of $n_i$'s in queue. Because of our assumption that $n_i$ is in the $S$ sets of exactly $D$ other nodes, this scenario would account for an $O(D)$ delay. However, such a chain is not the longest that may occur, as we see next. Suppose that $n_i$ and $n_k$ are two nodes in $S_j$, and in addition that a wait chain exists starting at $n_k$ and ending at $n_i$. In this wait chain, $n_k$ is waiting for a granted from the next node in the chain, which is waiting for a granted from the next node, which in turn is waiting for a release, and so on, all the way to the node that precedes $n_i$ in the chain, which is waiting for a granted from $n_i$. By Theorem 8.2, the algorithm is deadlock-free, so that this chain cannot involve more than $n - 1$ nodes. If the granted that sent to the node that is waiting on it on the chain is sent after $n_i$ receives the request from $n_i$, we see that a chain of messages of length $O(n)$ exists between a request sent by $n_i$ and its sending of release's. This is then the algorithm's time complexity, therefore the same as in the previous case.

As one last remark in this section, we encourage the reader to pursue the exercise of modifying Algorithm $A_{Mutually\_Exclude}$ for the case in which node $n_i$ is allowed to belong to $S_i$ (cf. Exercise 3).

8.2 Sharing multiple resources

Henceforth in this chapter, we no longer assume that only one single resource (or a group of resources that have to be accessed as a single entity) is shared by the nodes, but instead consider the more general case in which nodes may require access to resources in groups of varied composition. One immediate consequence of this relaxed view is that it is now possible for more than one node to be accessing shared resources concurrently, so long as no resource belongs to the group of resources accessed by any node.

Let $R$ be a set of resources, and let $R = |R|$. The members of $R$ are the resources $\rho_1, \ldots, \rho_R$, and for $1 \leq r \leq \rho_R$ with each resource $\rho_r \in R$ a set of nodes $S_r \subseteq N$ is associated. Nodes in
Section 8.3

The resource-sharing problem that we treat in this section is a generalization of the following paradigmatic problem, called the dining philosophers problem. Five philosophers sit at a round table, and five forks are placed on the table so that there is one fork between the plates of every two adjacent philosophers. A philosopher requires both forks that are adjacent to him in order to eat, and then it is impossible for neighbor philosophers to eat concurrently. The life of a philosopher is a cycle of thinking and eating for alternating finite periods of time. A solution to the problem consists of an algorithm that ensures mutual exclusion (a fork may only be used by one philosopher at a time), prevents deadlocks (at least one hungry philosopher must be eating), and prevents starvation (every philosopher must get to eat within a finite time of becoming hungry; incidentally, it is from the context of this problem that the term "starvation" in resource sharing comes).

In terms of our modeling of the previous section, in the dining philosophers problem \( N \) is the set of philosophers with \( n = 5 \), \( R \) is the set of forks with \( R = 5 \), and every one of \( S_1 \) through \( S_5 \) includes two philosophers that sit next to each other at the table (conversely, each philosopher is a member of exactly two such sets, specifically those that correspond to the forks that are adjacent to him). The graph \( G \) is then a five-node ring in which an edge corresponds to a fork.

In the generalized form of this problem, \( G \) is any connected undirected graph with one philosopher per node and one fork per edge. In order to eat, a philosopher must acquire all the forks that are adjacent to him. It is very important for the reader to note that the dining philosophers problem in this generalized form is entirely equivalent to the resource sharing problem, described in the previous section, in which a node always accesses the same set of resources. Although there is in principle no correspondence between the forks that are adjacent to a philosopher and those resources, a fork can be used to represent all the resources that two neighboring nodes share. Acquiring every adjacent fork is then equivalent to securing mutual exclusion in the access to all the resources a node needs. It is then to this generalized form of the dining philosophers problem that we dedicate the remainder of Section 8.3. Our discussion proceeds in two parts. In the first part, presented in Section
8.3.1, we give an algorithm to solve this generalized formulation of the problem. The second part, in Section 8.3.2, is dedicated to the extreme situation in which the thinking period of philosophers is negligibly small. Under this situation of perennial hunger, interesting issues appear related to the concurrency that can be achieved in the sharing of forks by the philosophers.

8.3.1 An algorithm

The solution to the generalized dining philosophers problem that we discuss in this section is given as Algorithm $A_{\text{Dine}}$. In this algorithm, node $n_i$ employs a Boolean variable, called $\text{hungry}_i$, and initialized to $\text{false}$, to indicate the need to access the resources that it shares with its neighbors. Whenever this variable becomes $\text{true}$, $n_i$ employs request messages to ask its neighbors to send it the forks it still does not have. Upon acquiring the forks corresponding to all the edges incident to it, $n_i$ accesses the shared resources and, perhaps contrary to our intuitive expectation, it does not send the forks to the corresponding neighbors, but rather keeps them to be sent when they are requested. However, if forks are indiscriminately distributed among the nodes at the beginning, deadlocks may, as it takes little effort to realize, occur. In addition, in the absence of some sort of priority among the nodes, the simple sending of forks upon receiving request's may easily lead to starvation.

The priority scheme that we adopt employs what we call a “turn” object per edge of $G$, much like forks. In addition, the turn associated with an edge can only be possessed by one of the nodes to which the edge is incident at a time, much like forks as well. What distinguishes turns from forks is that a node does not need to acquire turns for all the edges incident to it in order to access the shared resources, and also that turns get sent over the edges to a node's neighbors as soon as the node is through with accessing the shared resources. The essential goal of an edge's turn is to indicate which of its end nodes has the priority to hold that edge's fork when there is conflict. However, in the absence of conflict, that fork may be held by either node, even against the current location of the turn. Sending a turn over to the corresponding neighbor is a guarantee that the priority to hold that edge's fork alternates between the two nodes. In Algorithm $A_{\text{Dine}}$, a fork message is used to send a fork, while a turn message is used to send a turn. For all $n_j \in \text{Neig}_i$, node $n_i$ maintains two Boolean variables to indicate whether $n_i$ holds the fork and the turn that it shares with $n_j$. These are, respectively, $\text{holds}_{\text{fork}}^j$ and $\text{holds}_{\text{turn}}^j$, and should be initialized so that consistency is maintained over the edge $(n_i, n_j)$ (that is, exactly one of $n_i$ and $n_j$ must hold the fork and exactly one of them, not necessarily the same, must hold the turn). We do not provide these initial values when presenting the algorithm, but return to the issue later with a more detailed discussion.

When $n_i$ sends a $\text{request}$ and this message finds $\text{hungry}_i = \text{true}$ and $\text{false}$ upon arrival, $n_i$ does not send the corresponding fork over to $n_j$, but rather postpones the sending of this fork to after it has accessed the shared resource. If, on the other hand, the $\text{request}$ message finds $\text{hungry}_i = \text{true}$ and $\text{false}$ upon arrival, then the fork is sent at once to $n_j$, but $n_i$ must know that the fork is to be returned when it has completed its access to the shared resource. In order not to have to send two messages in each case (a turn and a fork in the former case, a fork and a request in the latter), every fork message is sent with a parameter, either as fork(nil), if all that needs to be achieved is the sending of a fork, or as fork(turn), to indicate that a turn is also being sent, or yet as fork(request), when a request for the fork is implied. For all $n_j \in \text{Neig}_i$, node $n_i$ employs the additional Boolean variable $\text{owes}_{\text{fork}}^j$, initially set to $\text{false}$, to remind it to send $n_j$ a fork when it finishes accessing the shared resources.
We assume that G's edges are FIFO, so that a request never overruns a fork or a turn, thereby ensuring that the arrival of a request at \( n_i \) from \( n_j \) only occurs when \( \text{owe}_i^j \) is true.

**Algorithm A_Dine:**

**Variables:**
- \( \text{hungry}_i = \text{false} \)
- \( \text{holds}_i^j \) for all \( n_j \in \text{Neig}_i \)
- \( \text{holds}_i^j \) for all \( n_j \in \text{Neig}_i \)
- \( \text{owe}_i^j = \text{false} \) for all \( n_j \in \text{Neig}_i \)

**Listing 8.11**

**Input:**
- \( \text{msg}_i = \text{nil} \).

**Action when not hungry and access to shared resources is needed:**
- \( \text{hungry}_i = \text{true} \);

Send request to all \( n_j \in \text{Neig}_i \) such that \( \text{hold}_i^j = \text{false} \).

**Listing 8.12**

**Input:**
- \( \text{msg}_i = \text{request} \) such that \( \text{origin}(\text{msg}_i) = (n_i, n_j) \).

**Action:**

if not hungry or not
begin

- \( \text{holds}_i^j := \text{false} \);

if not hungry then
- Send fork(nil) to \( n_j \)
else
- Send fork(request) to \( n_j \)
end
else

\( \text{owe}_i^j := \text{true} \).
Listing 8.13

Input:
\[ \text{msg}_i = \text{fork}(t) \text{ such that } \text{origin}(\text{msg}_i) = (n_i, n_j). \]

Action:
\[
\begin{align*}
\text{holds}_{-}\text{fork}_i^j & := \text{true}; \\
\text{if } t = \text{turn} \text{ then } \\
\text{holds}_{-}\text{turn}_i^j & := \text{true}; \\
\text{if } t = \text{request} \text{ then } \\
\text{owes}_{-}\text{fork}_i^j & := \text{true}; \\
\text{if } \text{for all } n_k \in \text{Neig} \text{ then } & \text{begin } \\
& \text{Access shared resources}; \\
& \text{hungry}_i := \text{false}; \\
& \text{for all } n_k \in \text{Neig} \text{ do } \\
& \text{if } \text{holds}_{-}\text{turn}_i^j \text{ then } & \text{begin } \\
& \text{holds}_{-}\text{turn}_i^j & := \text{false}; \\
& \text{if } \text{owes}_{-}\text{fork}_i^k \text{ then } & \text{begin } \\
& \text{owes}_{-}\text{fork}_i^k & := \text{false}; \\
& \text{holds}_{-}\text{fork}_i^k & := \text{false}; \\
& \text{Send } \text{fork}(\text{turn}) \text{ to } n_k & \text{end } \\
& \text{else } & \text{Send } \text{turn} \text{ to } n_k \\
& \text{end } \\
& \text{end. } \\
\end{align*}
\]

Listing 8.14

Input:
\[ \text{msg}_i = \text{turn} \text{ such that } \text{origin}(\text{msg}_i) = (n_i, n_j). \]

Action:
In Algorithm $A_{Dine}$, as in previous occasions in this chapter, the set $N_0$ comprises nodes for which the need to access the shared resources arises concurrently. Multiple executions of the algorithm coexist, and no variables are re-initialized for executions other than the very first. In this algorithm, actions $(8.11)$ through $(8.14)$ realize, respectively, the sending of request's for forks when the need arises for a node to access the shared resources, and the handling of request, fork, and turn messages.

In order to discuss the algorithm's main properties, we must at last be more specific about the initial values to be assigned to the $holds$$_fork$ and $holds$$_turn$ variables. Whereas any consistent assignment of values to the $holds$$_fork$ variables will do, the algorithm's properties are quite sensitive to the values that are assigned to the $holds$$_turn$ variables, and simple consistency across edges will not do in general.

Before continuing with this discussion on initial values, let us pause and introduce the concept of an orientation of the undirected graph $G$. As in Sections 6.3 and 7.2, such an orientation is a means of regarding $G$ as a directed graph, without however sacrificing the ability for messages to traverse $G$'s edges in both directions. In the present case, as in Section 6.3, $G$'s orientation will change dynamically, thereby increasing the importance of assigning directions to edges in a manner that is not too inflexible as assuming that $G$ is directed in the first place. An orientation of $G$ is a function

$$\omega : E \to N$$

such that, for $(n_i, n_j) \in E$, $\omega((n_i, n_j))$ is either $n_i$ or $n_j$, indicating respectively that, according to $\omega$, $(n_i, n_j)$ is directed from $n_i$ to $n_j$ or from $n_j$ to $n_i$. An orientation is said to be acyclic if it does not induce any directed cycle in $G$.

The following is then how the assignment of turns to neighbors is performed, and consequently the consistent assignment of values to the $holds$$_turn$ variables. Say that the edge $(n_i, n_j)$ is directed from $n_i$ to $n_j$ if the turn that corresponds to it is given to $n_j$. The initial assignment that we adopt is then such that the resulting orientation is acyclic. The essential importance of such an initial acyclic orientation is that, as the orientation changes by the sending of turns when a node is done with accessing the shared resources in $(8.13)$, that node becomes a source in the new orientation, that is, a node with all incident edges directed away from it. The resulting orientation is then acyclic as well, because the only changes in an orientation correspond to nodes that become sources, and then directed cycles that might have been formed would have to go through those nodes, which is impossible.

We now turn to a more formal statement of the algorithm's properties.

**Theorem 8.3.**

Algorithm $A_{Dine}$ ensures mutual exclusion in the access to the shared resources, and is in addition deadlock- and starvation-free.

**Proof:** By $(8.13)$, a node only accesses the shared resources if it holds the forks corresponding to all the edges incident to it. By $(8.12)$ and $(8.13)$, only one of every two neighbors may hold the fork that they share in any global state, and then no two neighbors can access the shared resources concurrently. Because by construction of $G$ nodes that are not neighbors never share any resources, mutual exclusion is guaranteed.

$G$'s orientation is always acyclic, and then $G$ always has at least one sink. Sinks are nodes that hold the turns corresponding to all the edges incident to them, and then by $(8.11)$ and $(8.12)$ must acquire all the forks that they do not hold within a finite time of having responded to the need to access the shared resources. No deadlock is then possible.

A node that is not a sink but does execute $(8.11)$ in order to access the shared resources is also guaranteed to acquire all the necessary forks within a finite time, and then no starvation ever occurs either. What supports this conclusion is that either such a node acquires all the
forks because its neighbors that hold turns do not need to access the shared resources (by (8.12) and (8.13)), or because it eventually acquires all the turns (by (8.13) and (8.14)) and then the forks as a consequence of the acyclicity of G's orientations.

The number of messages that need to be exchanged per access to the shared resources can be computed as follows. First a node may send as many request messages as it has neighbors, that is |Neig| in the case of node n. The worst that can happen is that n does not hold any turns and the request's that it sends find nodes that do not need to access the shared resources and then send n forks. Because n does not hold any turns, it may happen that these forks may have to be returned as fork(request) messages if n receives at least one request from its neighbors before receiving the last fork. By (8.13), n will then eventually receive all these forks back, then access the shared resources, and then send turns out. If we let

\[ \Delta = \max (|\text{Neig}|) \]

then clearly the algorithm's message complexity per access to the shared resources is \( O(\Delta) \). Message lengths are constant, and then the algorithm's bit complexity is also of \( O(\Delta) \).

The time complexity of Algorithm A_Dine per access to the shared resources is related to the longest chain of messages beginning with the sending of request's by a node and ending with the reception by that node of the last fork message that it expects. Such a chain happens for a node that is a source in the current acyclic orientation when all nodes require access to shared resources. In this case, the directed distance from that node to the sinks may be as large as \( n-1 \), and then the time complexity that we seek is \( O(n) \). One situation in which this worst case may happen is that of a ring with a single sink.

All of Algorithm A_Dine's properties rely strongly on the assumption of an initial acyclic orientation for G. Determining this initial acyclic orientation constitutes an interesting problem by itself, and appears to require randomized techniques to be solved unless nodes can be assumed to have distinct identifications totally ordered by <, as in Section 8.1 and other occasions in the book. If such is the case, then the initial acyclic orientation that we need can be determined with \( O(m) \) messages and \( O(1) \) time as follows. Every node sends its identification to its neighbors. For each edge \((n_i, n_j)\), the turn stays initially with \( n_i \) if \( id_i < id_j \); it stays with \( n_j \) if \( id_j < id_i \). When compared with the algorithms of Section 8.1, the approach of assigning priorities based on a dynamically evolving acyclic orientation of G can be regarded as trading the nonconstant message length in those cases by an initial overhead to establish the initial acyclic orientation.

### 8.3.2 Operation under heavy loads

A heavy-load situation is characterized by the fact that nodes require access to shared resources continually. Clearly, under such circumstances Algorithm A_Dine may be simplified to a great extent. Specifically, a node may send forks to all of its neighbors immediately upon finishing accessing the shared resources. As a consequence, request and turn messages are no longer needed, provided it is the placement of forks, instead of turns, that gives G's orientation. The only variables that are needed at \( n \) are then

\[ \text{holds}_\text{fork}_i \] for all \( n_i \in \text{Neig}_i \). The simplified algorithm is presented next as

**Algorithm A_Dine_H ("H" for Heavy), with \( N_h \) being the set of nodes that are sinks (i.e., hold all forks corresponding to incident edges) initially.**

```
Variables:

\[ \text{holds}_\text{fork}_i \] for all \( n_i \in \text{Neig}_i \)
```

Listing 8.15
**Input:**
\[
msg_i = \text{nil}.
\]

**Action if** \(n_i \in N_0\):
Access shared resources;

\[
\text{holds}_{-}\text{fork}^j_i : = \text{false} \text{ for all } n_j \in \text{Neig}_i;
\]
Send fork to all \(n_j \in \text{Neig}_i\).

---

**Listing 8.16**

**Input:**
\[
msg = \text{fork such that origin}(msg) = (n_i, n_j).
\]

**Action:**

\[
\text{holds}_{-}\text{fork}^j_i : = \text{true};
\]

\[
\text{if for all } n_k \in \text{Neig} \text{ then begin}
\]
Access shared resources;

\[
\text{holds}_{-}\text{fork}^j_i : = \text{false} \text{ for all } n_k \in \text{Neig}_i;
\]
Send fork to all \(n_k \in \text{Neig}_i\).

end.

---

Actions (8.15) and (8.16) are both related to (8.13) of Algorithm \(A_{Dine}\). It should then come with no difficulty that Theorem 8.3 is equally applicable to Algorithm \(A_{Dine_H}\) as well. In the remainder of this section, we turn to the synchronous model of computation for a more detailed analysis of Algorithm \(A_{Dine_H}\). Our choice of a synchronous model for this analysis is motivated by the simplicity that ensues from that model, although corresponding results for the asynchronous model also exist and can be found in the literature.

The synchronous counterpart of Algorithm \(A_{Dine_H}\) starts off with an initial acyclic orientation at pulse \(s = 0\) and generates a sequence of acyclic orientations for pulses \(s > 0\). For \(s \geq 0\), at pulse \(s\) all sinks concurrently access shared resources and then send forks to neighbors. The evolution of acyclic orientations in synchronous time is such that a new acyclic orientation is generated by reversing the orientation of all edges incident to sinks, which then become sources in the new orientation. This mechanism is referred to as the edge-reversal mechanism, having applications beyond the context of resource sharing. For example, together with Algorithm \(A_{Schedule_AS}\) of Section 5.3.2, an algorithm based on the edge-reversal mechanism is of key importance as a technique for time-stepped simulation (cf. Section 10.2). Both in this section and in Section 10.2, the edge-reversal mechanism is employed to schedule nodes for operation so that neighbors do not operate concurrently. In this section, the "operation" is to access resources that nodes share with neighbors, while in Section 10.2 the term has a different meaning. Because of its role as a scheduler, it is also common to find the edge-reversal mechanism referred to as scheduling by edge reversal.

Let \(\omega_1, \omega_2, \ldots\) denote the sequence of acyclic orientations created by the edge-reversal mechanism, and \(\text{Sinks}_1, \text{Sinks}_2, \ldots\) denote the corresponding sets of sinks. For \(k \geq 1\), \(\omega_k\) is the
orientation at pulse $s = k - 1$. For $n_i \in N$ and $k \geq 1$, let $m_i(k)$ be the number of times $n_i$ appears in $Sinks_1, \ldots, Sinks_k$.

**Theorem 8.4.**
Consider two nodes $n_i$ and $n_j$, and let $r \geq 1$ be the number of edges on a shortest undirected path between them in $G$. Then $|m_i(k) - m_j(k)| \leq r$ for all $k \geq 1$.

**Proof:** We use induction on the number of edges on a shortest undirected path between $n_i$ and $n_j$. The case of one edge constitutes the basis of the induction, and then the assertion of the theorem holds trivially, as in this case $n_i$ and $n_j$ are neighbors in $G$, and must therefore appear in alternating sets in $Sinks_1, Sinks_2, \ldots$. As the induction hypothesis, assume the assertion of the theorem holds whenever a shortest undirected path between $n_i$ and $n_j$ has a number of edges no greater than $r - 1$. When $n_i$ and $n_j$ are separated by a shortest undirected path with $r$ edges, consider any node $n_\ell$ (other than $n_i$ and $n_j$) on this path and let $d$ be the number of edges between $n_i$ and $n_\ell$ on the path. By the induction hypothesis,

\[
|m_i(k) - m_\ell(k)| \leq d
\]

and

\[
|m_i(k) - m_j(k)| \leq r - d.
\]

yielding

\[
|m_i(k) - m_j(k)| \leq r.
\]

thence the theorem.

**Theorem 8.4** is in fact more than a mere starvation-freedom statement, as not only does it imply that all nodes become sinks within a bounded number of pulses, but it also establishes a bound on the relative frequency with which nodes become sinks. This theorem is then a stronger version of **Theorem 8.3** as far as starvation is concerned.

The number of distinct acyclic orientations of $G$ is of course finite, so the sequence $\omega_1, \omega_2, \ldots$ must at some pulse embark in a periodic repetition of orientations, which we call a periodic *orientation*, or simply a period (Figure 8.1). Orientations in a period are said to be periodic *orientations*. The next corollary establishes an important property of periods.

**Corollary 8.5.**
The number of times that a node becomes a sink in a period is the same for all nodes.
A period of five orientations results from the edge-reversal mechanism started at the orientation shown in the upper left corner of the figure, which is outside the period. In this period, every node becomes a sink twice.

**Proof:** Suppose, to the contrary, that two nodes \( n_i \) and \( n_j \) exist that become sinks different numbers of times in a period. Suppose, in addition, that a shortest undirected path between \( n_i \) and \( n_j \) in \( G \) has \( r \) edges. Letting \( p \) be the number of orientations in the period and \( k = (r + 1)p \) yields

\[
|m_i(k) - m_j(k)| \geq r + 1,
\]

which contradicts Theorem 8.4.

Clearly, the period is unequivocally determined given \( \omega_1 \). We let \( m(\omega_1) \) denote the number of times that nodes become sinks in this period, and \( p(\omega_1) \) denote the number of orientations in the same period.

One issue of great interest is the "amount of concurrency" that scheduling by edge reversal is capable to yield from an initial acyclic orientation \( \omega_1 \). The importance of this issue comes from the very nature of the resource-sharing computations we have been considering in all of
Section 8.3, and from the intuitive realization that the choice of $\omega_1$ greatly influences the number of nodes that become sinks concurrently. The measure that we adopt for the concurrency attainable from $\omega_1$, denoted by $\text{Conc}(\omega_1)$, is the average, over a large number of pulses and over $n$, of the number of times each node becomes a sink in those pulses. More formally, 
$$\text{Conc}(\omega_1) = \lim_{k \to \infty} \frac{1}{k} \sum_{n=0}^{k} m_n(k).$$

Clearly, we get more concurrency as more nodes become sinks earlier.

**Theorem 8.6.**
$$\text{Conc}(\omega_1) = m(\omega_1)p(\omega_1).$$

**Proof:** For some $\ell \geq 1$, let $\omega_1$ be the first periodic orientation in $\omega_1, \omega_2, \ldots$. For $k \geq \ell$, the first $k$ orientations of $\omega_1, \omega_2, \ldots$ include $(k - \ell + 1)/p(\omega_1)$ repetitions of the period, so
$$k = \left[ \frac{k - \ell + 1}{p(\omega_1)} \right] p(\omega_1) + u$$

and
$$\sum_{n_i \in N} m_n(k) = n \left[ \frac{k - \ell + 1}{p(\omega_1)} \right] m(\omega_1) + v,$$

where $\ell - 1 \leq \ell + p(\omega_1) - 2$ and $u \leq \ell \leq n$. The theorem then follows easily in the limit as $k \to \infty$.

It follows immediately from **Theorem 8.6** that
$$\frac{1}{n} \leq \text{Conc}(\omega_1) \leq \frac{1}{2}.$$ 
This is so because it takes at most $n$ pulses for a node to become a sink (the longest directed distance to a sink is $n - 1$), so
$$m(\omega_1) \geq \frac{p(\omega_1)}{n},$$

and because the most frequently that a node can become a sink is in every other pulse, so
$$m(\omega_1) \leq \frac{p(\omega_1)}{2}.$$ 

If $G$ is a tree, then it can be argued relatively simply that $\text{Conc}(\omega_1) = 1/2$, regardless of the initial orientation $\omega_1$. If $G$ is not a tree, then, interestingly, $\text{Conc}(\omega_1)$ can also be expressed in purely graph-theoretic terms, without recourse to the dynamics of the edge-reversal mechanism. For such, let $k$ denote an undirected cycle in $G$ (with $|k|$ nodes). Let also $n^+(k, \omega_1)$ and $n^-(k, \omega_1)$ denote the number of edges in $k$ oriented by $\omega_1$ clockwise and counterclockwise, respectively. Define
$$\rho(k, \omega_1) = \frac{1}{|k|} \min \{ n^+(k, \omega_1), n^-(k, \omega_1) \},$$

and let $K$ denote the set of all of $G$’s undirected cycles.

**Theorem 8.7.**
If $G$ is not a tree, then $\text{Conc}(\omega_1) = \min_{k \in K} \rho(k, \omega_1)$.

**Proof:** The proof is quite involved, and escapes the intended scope of this book. The interested reader is referred to the pertaining literature.

There are in the literature additional results concerning scheduling by edge reversal that we do not explicitly reproduce here. Some are positive, as the one that states that this mechanism is optimal (provides most concurrency) among all schemes that prevent neighbors from operating concurrently and require neighbors to operate alternately. Other
results are negative, as for example the computational intractability (NP-hardness) of finding the initial acyclic orientation $\omega_1$ that optimizes $\text{Conc}(\omega_1)$.

8.4 The drinking philosophers problem

If nodes may access different subsets of resources whenever they require access to shared resources, then the possibility that neighbors in $G$ access shared resources concurrently exists, provided the sets of resources they access have an empty intersection. In such cases, the technique of employing one single fork per edge to secure exclusive access to the resources that two neighbors share is no longer sufficient. Instead, associated with every edge there has to be one object for each resource that the corresponding neighbors share. Such objects are bottles from which the philosophers drink, and the problem of ensuring mutual exclusion, deadlock-freedom, and starvation-freedom in the drinking of the philosophers is referred to as the drinking philosophers problem.

At node $n$, the set of bottles shared over edge $(n_i, n_j)$ is denoted by $B_{i,j}^i$ (which, clearly, is the same as $B_{j,i}^j$). For all $n_i \in \text{Neig}_i$, and for $b_k \in B_i$, node $n_i$ employs the Boolean variable $\text{needs}_i^k$, initially set to false, to indicate whether $b_k \in B_i$ is needed. The need to access shared resources is indicated at $n_i$ by the Boolean variable $\text{thirsty}_i$, initialized to false. In order to access the shared resources that it requires, $n_i$ must acquire all the bottles corresponding to those resources, i.e., $b_k \in B_i$ such that $\text{needs}_i^k = \text{true}$ for all $n_i \in \text{Neig}_i$.

Except for this need to acquire multiple objects corresponding to a same edge, the solution that we describe for the drinking philosophers problem is quite similar to that of the dining philosophers problem. In particular, the same priority scheme based on the turn objects that we employed in Algorithm $A\text{-Dine}$ is also used in the new solution. In contrast with the multiplicity of bottles for each edge, the turns continue to exist in the number of one per edge, as in the dining philosophers case. What this amounts to is that turns are sent out on all edges when a node finishes accessing a group of shared resources, even in those edges whose bottles were not needed.

This solution to the drinking philosophers problem is given next as Algorithm $A\text{-Drink}$. Additional variables employed by node $n_i$ are, for all $n_j \in \text{Neig}_i$, the Boolean $\text{holds-turn}_i^j$ (employed as in the dining philosophers case), and for all $b_k \in B_i$ and for all $n_j \in \text{Neig}_i$, the Boolean $\text{holds-bottle}_i^k$, and $\text{owes-bottle}_i^k$ (this one initially set to false), employed respectively to indicate whether $n_i$ holds $b_k$ and whether $n_i$ has postponed the sending of $b_k$ to $n_j$ until it is through with accessing the shared resources. As in the dining philosophers case, the $\text{holds-turn}$ and $\text{holds-bottle}$ variables must be initialized to consistency across edges. In addition, the $\text{holds-turn}$ variables must be initialized as to induce an acyclic orientation on $G$. Auxiliary set variables $X$ and $Y$, both initially empty, are used by $n_i$.

The messages employed by Algorithm $A\text{-Drink}$ are completely analogous to those of Algorithm $A\text{-Dine}$, but with some slight modifications. A message $\text{request}(X)$ is sent by $n_j$ to request a set of bottles $X \subseteq B_{i,j}^j$. A set $X$ of bottles is sent via the message $\text{bottle}(X)$, which similarly to the fork message of the dining philosophers
case may carry an additional nil, turn, or request(Y) parameter. Finally, a turn message is used to send a turn over an edge.

**Algorithm A-Drink:**

Variables:

\[
\text{thirsty}_i = \text{false}; \\
\text{holds_{bottle}}_i \quad \text{for all } n_j \in \text{Neig}_i \text{ and all } b_k \in B_i^j; \\
\text{holds_{turn}}_i \quad \text{for all } n_j \in \text{Neig}_i; \\
\text{owes_{bottle}}_i \quad = \text{false} \text{ for all } n_j \in \text{Neig}_i \text{ and all } b_k \in B_i^j; \\
\text{needs_{bottle}}_i \quad = \text{false} \text{ for all } n_j \in \text{Neig}_i \text{ and all } b_k \in B_i^j; \\
X = \emptyset; \\
Y = \emptyset.
\]

**Listing 8.17**

Input:

\[\text{msg}_i = \text{nil}.\]

Action when not thirsty and access to shared resources is needed:

\[\text{thirsty}_i := \text{true}; \]

\[\text{needs_{bottle}}_i := \text{true} \text{ for all } n_j \in \text{Neig}_i \text{ and all } b_k \in B_i^j; \]

such that access to the resource for which \(b_k\) stands is needed;

\[\text{for all } n_j \in \text{Neig}_i \text{ such that there exists } b_k \in B_i^j \text{ with } \text{false} \] do

\[\text{begin} \]

Let \(X\) be the subset of \(B_i^j\) such that \(b_k \in X\) if and only if

\[\text{needs_{bottle}}_i := \text{true} \text{ and } \text{holds_{bottle}}_i := \text{false}; \]

\[\text{Send request (X) to } n_j; \]

\[X := \emptyset; \]

\[\text{end}. \]
Listing 8.18

Input: 
\[ \text{msg}_i = \text{request}(X) \text{ such that } \text{origin}(	ext{msg}_i) = (n_i, n_j). \]

Action:
\[
\text{for all } b_k \in X \text{ do } \\
\quad \text{if not } \text{thirsty}_i \text{ or not } \\
\quad \quad \text{begin } \\
\quad \quad \quad \text{holds}_i \text{ := false; } \\
\quad \quad \quad X_i := X_i \cup \{b_k\}; \\
\quad \quad \quad \text{if } \text{thirsty}_i \text{ and } \\
\quad \quad \quad \quad \text{needs}_i \text{ := Y}_i \cup \{b_k\} \\
\quad \quad \text{end} \\
\quad \text{else } \\
\quad \quad \text{owes}_i \text{ := true; } \\
\quad \text{if } X_i \neq \emptyset \text{ then } \\
\quad \quad \text{begin } \\
\quad \quad \quad \text{if } Y_i = \emptyset \text{ then } \\
\quad \quad \quad \quad \text{Send bottle } (X_i, \text{nil}) \text{ to } n_j \\
\quad \quad \quad \text{else } \\
\quad \quad \quad \quad \text{begin } \\
\quad \quad \quad \quad \quad \text{Send bottle } (X_i, \text{request}(Y_i)) \text{ to } n_j; \\
\quad \quad \quad \quad \quad Y_i := \emptyset; \\
\quad \quad \quad \quad \text{end; } \\
\quad \quad \quad X_i := \emptyset; \\
\quad \quad \text{end. } \\
\]

Listing 8.19

Input: 
\[ \text{msg}_i = \text{bottle}(X, t) \text{ such that } \text{origin}(	ext{msg}_i) = (n_i, n_j). \]

Action:
\[
\text{holds}_i \text{ := true for all } b_k \in X; \\
\text{if } t = \text{turn} \text{ then } \\
\]

\[
\text{holds\_turn}_i^j := \text{true};
\]
if \( t = \text{request}(Y) \) then
\[
\text{owes\_bottle}_i^{jk} := \text{true} \text{ for all } b_k \in Y;
\]
if \( \text{holds\_bottle}_i^{kl} \) then
begin
Access shared resources;

\[
\text{thirsty} := \text{false};
\]
for all \( n_k \in \text{Neig} \) and all \( b_\ell \in b_i^k \) then
begin
\[
\text{thirsty} := \text{false};
\]
for all \( n_k \in \text{Neig} \) do
if \( \text{holds\_turn}_i^j \) then
begin
\[
\text{holds\_turn}_i^j := \text{false};
\]
for all \( b_\ell \in b_i^k \) do
if \( \text{holds\_bottle}_i^{kl} \) then
begin
\[
\text{owes\_bottle}_i^k := \text{false};
\]
\[
\text{holds\_bottle}_i^{kl} := \text{false};
\]
\[
X_i := X_i \cup \{ b_\ell \};
\]
end;
if \( X_i \neq \emptyset \) then
begin
Send bottle \((X_i, \text{turn})\) to \( n_i \);
\[
X_i := \emptyset;
\]
end
else
Send \text{turn} to \( n_i \)
end.
end.

Listing 8.20

Input:
\( \text{msg}_i = \text{turn} \) such that \( \text{origin}_i(\text{msg}_i) = (n_i, n_i) \).
Action:
\[
\text{holds\_turn}_i^j := \text{true}.
\]
Actions (8.17) through (8.20) are entirely analogous to actions (8.11) through (8.14), respectively, of Algorithm A-Dine. Because the same priority scheme is used in both algorithms, Theorem 8.3 is, in essence, applicable to Algorithm A-Drink as well. With the exception of the bit complexity, the two algorithms also share the same complexity measures. The bit complexity of Algorithm A-Drink is different because request and bottle messages carry references to a set of bottles (possibly two sets of bottles, in the case of bottle messages) with a nonconstant number of bottles. Because two neighbors share as many bottles as they share resources, and because they share at most $R$ resources (cf. Section 8.2), the algorithm’s bit complexity is $O(\Delta R \log R)$, provided every resource can be identified with $\lceil \log R \rceil$ bits.

8.5 Exercises

1. In Algorithm A_Mutually-Exclude-C, show that the sequence numbers can be implemented with $O(\log n)$ bits.
2. Show, for Algorithm A-Mutually-Exclude, that it may be necessary to compare two sequence numbers differing from each other by an arbitrarily large amount.
3. Modify Algorithm A-Mutually-Exclude for the case in which node $n_i$ may belong to $S_i$.
4. In the context of Section 1.5, find the $r(c)$’s for Algorithm A-Dine-H.

8.6 Bibliographic notes


The formulation of Section 8.2 can be found in Barbosa (1986) and in Barbosa and Gafni (1987, 1989b). The dining philosophers problem appeared originally in Dijkstra (1968), and received attention in a distributed setting, either as posed originally or as variations thereof, in Chang (1980), Lynch (1980), Lynch (1981), and Rabin and Lehmann (1981). Algorithm A-Dine of Section 8.3.1 is based on Chandy and Misra (1984), but the idea of transforming sinks into sources to maintain the acyclicity of a graph’s orientation appeared previously (Gafni and Bertsekas, 1981) in the context of routing in computer networks. The analysis that appears in Section 8.3.2 for the heavy-load case is from Barbosa (1986) and Barbosa and Gafni (1987, 1989b), where the omitted proof of Theorem 8.7 also appears, as well as other results related to optimality and intractability (in the sense of NP-hardness, as in Karp (1972) and Garey and Johnson (1979)). Most of the concurrency notions involved with scheduling by edge reversal are closely related to the concept of a multicoloring of a graph’s nodes. Such a concept can be looked up in Stahl (1976), for example. Other sources of information on scheduling by edge reversal are Bertsekas and Tsitsiklis (1989), Malka, Moran, and Zaks
Chapter 9: Program Debugging

Overview

Debugging is the part of the program development process whereby conceptual and programming errors are detected and corrected. The debugging of a sequential program is achieved mainly through the use of rather simple techniques that involve the ability to re-execute a program and to halt its execution at certain points of interest (the so-called breakpoints). Asynchronous algorithms like the ones we have been treating in this book lack both the determinism that makes the re-execution of sequential programs simple, and the unique total order of events that facilitates the detection of states where a halt is desired. Clearly, then, the debugging of programs based on such algorithms is altogether a different matter.

Notwithstanding this difference in levels of difficulty, approaches to the debugging of programs based on asynchronous algorithms have concentrated on the same two major techniques on which the debugging of sequential programs is based, namely deterministic re-execution and breakpoint detection. It is then to these two major topics that we dedicate this chapter, beginning in Section 9.1 with some preliminary concepts, and then progressing through Sections 9.2 and 9.3, respectively on techniques for program re-execution and breakpoint detection. Throughout the chapter, \( G \) is an undirected graph.

The detection of breakpoints can be an especially intricate endeavor, depending on the characteristics of the breakpoint one is seeking to detect. For this reason, in Section 9.3 we limit ourselves to very special classes of breakpoints, chiefly those that are either unconditional or depend on predicates that can be expressed as logical disjunctions or conjunctions of local predicates. In this context, we provide techniques that fall into two classes, specifically those that are based on a re-execution (in the style of Section 9.2), and those that are not so.

Sections 9.4 and 9.5 contain, respectively, exercises and bibliographic notes.

9.1 Preliminaries

The debugging of a sequential program is a cyclic process supported by two basic techniques, those of program re-execution and of breakpoint detection. We assume henceforth that the programs that we treat in this chapter, sequential and distributed alike, never act on probabilistic decisions, and then a sequential program is guaranteed to go through the same sequence of states whenever it is re-executed from the same initial conditions. In the asynchronous distributed case, however, there are sources of nondeterminism other than those related to probabilistic decisions, specifically those related to the model's unsynchronized local clocks and unpredictable delays for message delivery among neighbors. As a consequence, the simple re-execution from the same initial conditions is not enough to ensure that all nodes will repeat the same behavior as in the previous execution.

This issue of nondeterminacy is also what distinguishes the sequential and asynchronous distributed cases when it comes to the detection of breakpoints during an execution. In the sequential case, all operations on variables are totally ordered, and then checking for the occurrence of particular states where predefined predicates hold poses no conceptual difficulties. In the asynchronous distributed case, on the other hand, no such unique total order exists, and the detection of global states with the characteristics required by a predefined predicate becomes a much harder problem.

The key to approaching the two problems is the treatment of timing issues under the asynchronous model that we pursued in Section 3.1. Specifically, in order to reproduce an execution of an asynchronous algorithm, it suffices to ensure that the re-execution follows the exact same partial order of events that was generated by the original execution.
Detecting breakpoints correctly is also very much dependent upon the concepts introduced in that section, because, as we already mentioned, what is required is the detection of global states at which the required predicates hold. However, what is needed is not just an algorithm like Algorithm A_Record_Global_State of Section 5.2.1, which offers no control as to which global state it records, but rather algorithms that are guaranteed not to miss a global state with the desired characteristics if one exists.

Before we proceed to the remaining sections of the chapter, let us pause briefly for a few terminological comments. Although in these first two sections we have attempted to comply with the standard practice of reserving the terms "algorithm" and "program" for different entities (a program is normally a realization of an algorithm, involving a particular programming language and often assumptions on the system's architecture), henceforth we shall drop the distinction and refer to the debugging of an algorithm as encompassing the debugging of programs as well (even though what is normally true is the converse). We do this to simplify the terminology only, and no further presumptions are implied.

9.2 Techniques for program re-execution

As we remarked in the previous section, the aim when attempting to re-execute an asynchronous algorithm is to re-generate the same set of events, and hence the same partial order among them, as in the original execution. What is needed to achieve this goal is twofold. First of all, the set $N_0$ of spontaneous initiators must be the same as in the original execution. Secondly, from the perspective of individual nodes, messages must be received in the same order as in the original execution. Note that this involves more than the order of message reception on a particular edge, as in reality what is required is the preservation of the order of message reception across all edges incident to a node. If edges are FIFO, then it suffices that a node, during the re-execution, consider the edges to receive messages in the same order as they happened to be considered in the original execution. Put differently, if node $n$ has as neighbors nodes $n_j, \ldots, n_k$ for $k = |Neig|$, and in the original execution the first message was received at $n$ from $n_a \in \{n_j, \ldots, n_k\}$, the second message from $n_b \in \{n_j, \ldots, n_k\}$, and so on, then during the re-execution the same order must be respected. The solution that we describe next is given for this case of FIFO edges, although an extension to the case in which edges are not FIFO can also be devised (cf. Exercise 1).

Preserving this order of appearance of a node's neighbors (equivalently, of a node's incident edges) in the sequence of messages the node receives during the re-execution of an algorithm can be achieved through the following two-phase process. During the original execution, every node records a sequence of pointers to neighbors; this recording is the process's first phase. The second phase occurs during the re-execution, in which the sequences recorded during the first phase are employed to force nodes to receive messages from neighbors in the order implied by the recorded sequence. The combined sequences recorded by all nodes during the first phase constitute a trace of the original execution. This trace, along with the composition of the $N_0$ set associated with the original execution, clearly suffice for the algorithm to be deterministically re-executed.

Before we proceed to describe more precisely the mechanism whereby the trace is employed during the re-execution, it must be mentioned that the recording of the trace may cause the original execution to be different from what it would be if no recording were being done. This is the so-called probe effect of the tracing recording process, and is of little consequence from a purely theoretical point of view, because under the assumptions of the asynchronous model any execution is as good as any other. However, in a practical setting where certain executions would be favored by certain prevailing timing conditions, the probe effect can be misleading, in the sense of causing the trace of an "atypical" execution to be recorded. In such circumstances, probe effects are important and the recording of traces (i.e., the "probe") should be designed to keep them to a minimum.

During the trace recording, node $n$ employs the Boolean variable $\text{initiator}_n$, initialized to $\text{false}$, to indicate whether $n$ turns out to be a member of $N_0$. In addition, a queue
of pointers to neighbors is employed by \( n_i \) to record the origins of all the messages it receives during the execution. This queue is called \( queue \), and is initialized to \( \text{nil} \). At all times, \( \text{first\_in\_queue}_i \) is assumed to be the first element in \( \text{queue}_i \), being equal to \( \text{nil} \) if \( \text{queue}_i = \text{nil} \). During the re-execution phase, the Boolean \( \text{initiator}_i \) is used to make up an \( N_0 \) set that is equal to the one of the original execution. Similarly, \( \text{queue}_i \) is employed to control the reception of messages by node \( n_i \). This is achieved by conditioning actions on the reception of a message \( \text{msg}_i \) to be executed only when \( \text{origin}_i(\text{msg}_i) = (n_i, n_j) \), provided \( \text{queue}_i \) is updated by the removal of its first element whenever a message is received. The mechanism whereby this takes place is through the use of the Boolean conditions allowed in our general template, Algorithm \( \text{A\_Template} \) of Section 2.1.

Next we present two asynchronous algorithms, one for each of the phases involved in the deterministic re-execution of asynchronous algorithms. These algorithms are called Algorithm \( \text{A\_Record\_Trace} \) and Algorithm \( \text{A\_Replay} \), respectively for the first phase and the second phase. These two algorithms are derived directly from Algorithm \( \text{A\_Template} \), and are therefore intended for generic asynchronous computations that fit that template.

### Algorithm \( \text{A\_Record\_Trace} \):

**Variables:**
- \( \text{Initiator}_i \) = false;
- \( \text{queue}_i \) = \( \text{nil} \);
- \( \text{first\_in\_queue}_i \) = \( \text{nil} \);
- Other variables used by \( n_i \) and their intial values, are listed here.

**Listing 9.1**

**Input:**
- \( \text{msg}_i = \text{nil} \).

**Action if** \( n_i \in N_0 \):
- \( \text{initiator}_i \) := true;
- Do some computation;
- Send one message on each edge of a (possibly empty) subset of \( \text{Inc}_i \).

**Listing 9.2**

**Input:**
- \( \text{msg}_i \) such that \( \text{origin}_i(\text{msg}_i) = (n_i, n_j) \).

**Action:**
- Append \( n_j \) to \( \text{queue}_i \);
- Do some computation;
- Send one message on each edge of a (possibly empty) subset of \( \text{Inc}_i \).

In Algorithm \( \text{A\_Record\_Trace} \), the variable \( \text{initiator}_i \) is set to true in (9.1) to signal to the re-execution phase what the members of \( N_0 \) must be. Similarly, the portion of
the trace that corresponds to message receptions at \( n_i \) is recorded in \( queue_i \) in (9.2). In Algorithm A_Replay, presented next, the members of \( N_0 \) are exactly those nodes \( n_i \) for which \( \text{initiator}_i = \text{true} \) at the end of the execution of Algorithm A_Record_Trace. In this sense, \( N_0 \) is no longer a set of spontaneous initiators, but rather the set of nodes that are forced to initiate the computation so the original execution of the algorithm can be faithfully reproduced.

**Algorithm A_Replay:**

<table>
<thead>
<tr>
<th>Variables:</th>
</tr>
</thead>
<tbody>
<tr>
<td>( queue_i );</td>
</tr>
<tr>
<td>( \text{first}<em>{\text{in}}</em>{\text{queue}}_i );</td>
</tr>
</tbody>
</table>

Other variables used by \( n_i \) and their initial values, are listed here.

**Listing 9.3**

<table>
<thead>
<tr>
<th>Input:</th>
</tr>
</thead>
<tbody>
<tr>
<td>( msg_i = \text{nil} ).</td>
</tr>
</tbody>
</table>

**Action if \( n_i \in N_0 \):**

- Do some computation;
- Send one message on each edge of a (possibly empty) subset of \( \text{Inc}_i \).

**Listing 9.4**

<table>
<thead>
<tr>
<th>Input:</th>
</tr>
</thead>
<tbody>
<tr>
<td>( msg_i ) such that ( \text{origin}(msg_i) = (n_i, n_j) ).</td>
</tr>
</tbody>
</table>

**Action when \( n_j = \text{first}_{\text{in}}_{\text{queue}}_i \):**

- Remove \( \text{first}_{\text{in}}_{\text{queue}}_i \) from \( queue_i \);
- Do some computation;
- Send one message on each edge of a (possibly empty) subset of \( \text{Inc}_i \).

Action (9.3) is executed by the nodes that initiate the re-execution according to the trace recorded by Algorithm A_Record_Trace. Action (9.4), which is only executed by \( n_i \) on a message arriving on edge \((n_i, n_j)\) when \( n_j = \text{first}_{\text{in}}_{\text{queue}}_i \), ensures that message receptions are acted upon by \( n_i \) in the same order as they were in the execution of Algorithm A_Record_Trace.

As a final remark in this section, note that the message, time, and bit complexities of both Algorithm A_Record_Trace and Algorithm A_Replay are the same. This is only expected, in view of the correctness of Algorithm A_Replay in reproducing the execution of Algorithm A_Record_Trace, if we consider that neither the recording process nor the deterministic re-execution employ any messages in addition to those already present in the original computation.

**9.3 Breakpoint detection**

This section is devoted to the second major problem we discuss in this chapter in connection with the debugging of asynchronous algorithms, namely that of detecting breakpoints. Our discussion proceeds in Sections 9.3.1 through 9.3.3 as follows. Section 9.3.1 contains fundamental definitions and concepts, especially those related to the types of breakpoints to be treated in the sequel. The remaining two sections are devoted each to a different aspect
of the problem. **Section 9.3.2** contains a trace-based approach, and **Section 9.3.3** an approach that does not depend on a trace of a previous execution. The bulk of **Section 9.3** is contained in **Section 9.3.3**, where we introduce a collection of distributed algorithms for the detection of some types of breakpoints. Although **Section 9.3.2** also contains interesting insights into the problem, it is on **Section 9.3.3** that the reader should concentrate.

### 9.3.1 Fundamentals

A breakpoint in the execution of an asynchronous algorithm is a global state at which one wishes the computation to halt so that nodes' variables can be examined. All the breakpoints we study in the forthcoming sections refer to local states of the nodes only, so that for our purposes in this chapter messages in transit on edges are an unimportant part of a global state. Whenever the need arises for a global state to be represented, we shall then do so by considering nodes' local states only. For node \( n \), we let \( lt \geq 0 \) be \( n \)'s local time. For simplicity when describing our algorithms, we assume that \( lt \) is in fact an event counter at \( n \), that is, \( lt \) = 0 initially and is increased by one either upon the spontaneous sending of messages by \( n \), if \( n \in Neig \), or upon the reception of a message by \( n \). Clearly, a node \( n \)'s local state is unequivocally determined given \( lt \).

Because messages in transit play no role in the global states of our interest, and considering the relationship we just described between a node's local state and local time, a node's view of a global state can be represented by an \( n \)-component array of local times. In such an array, and for \( 1 \leq i \leq n \), the \( i \)th component contains some value of \( lt \). If we revisit our initial study of global states in **Section 3.1**, then clearly an \( n \)-component array \( \varphi \) of natural numbers is a global state if and only if no \( n \in N \) exists that ever receives a message earlier than (or at) \( \varphi [j] \) that was sent by some \( n \in Neig \), later than \( \varphi [j] \). The definition of an earliest global state with respect to some property that we alluded to in **Section 3.1** can in this simplified view of a global state be given as follows. A global state \( \varphi \) is the earliest global state for which a certain property holds if and only if no other global state \( \varphi' \) for which the property also holds is such that \( \varphi [j] \leq \varphi' [j] \) for all \( n \in N \). Depending on the particular property one is considering, it is conceivable, as we will see later in **Section 9.3.3**, that more than one earliest global state exists. In this case, all the earliest global states are incomparable to one another, in the sense that the past of none of them is in the past of another (or, equivalently, if we resort to the terminology of **Section 3.1**, none of them comes earlier than any other).

There is a great variety of problems that may be considered when studying breakpoints of asynchronous algorithms, so right at this introductory section it must remain very clear which problems that we consider are. The first important distinction is related to the so-called "weak" and "strong" senses in which breakpoints can be treated. The weak sense refers to breakpoints as global states of one single execution of the algorithm, while the strong sense is about breakpoints in all possible executions of the algorithm. Although attempts are described in the literature that focus on strong-sense problems, it is to be intuitively expected that such attempts invariably result in computationally intractable problems for general computations, given the prohibitively large number of possible executions of an asynchronous algorithm. It is then to problems in the weak sense that we dedicate our study. One important consequence of restricting ourselves to such problems is that, if a particular execution of an asynchronous algorithm fails to contain a global state with certain desired properties, one cannot infer that no execution exists in which such a global state would appear.

Although in many situations the ultimate goal of considering breakpoints in asynchronous algorithms is to halt the execution at the corresponding global state, with one single exception it is not to this halting problem that we dedicate most of our efforts in the forthcoming sections within **Section 9.3**, but rather to the problem of only detecting the occurrence of the breakpoints. The exception is the material that we present in **Section 9.3.2**.
where the halting problem is considered. In all other situations, that is, those discussed in Section 9.3.3, if halting is desired after the detection, then special techniques of the so-called "checkpointing and rollback recovery" type must be employed so that the execution can be "returned" to the global state where the breakpoint was detected. We pursue the issue no further in this book, but in Section 10.4 the reader can find closely related techniques, only in a totally different context.

A breakpoint can be unconditional or conditional. An unconditional breakpoint is specified by providing a local unconditional breakpoint for each \( n \in N \), denoted by \( lub \). The local unconditional breakpoint of \( n \) is either a nonnegative integer specifying the value of \( lt \) with which \( n \) is to participate in the breakpoint, or such that \( lub = \infty \) if \( n \) does not participate in the breakpoint. This flexibility of allowing nodes not to participate in breakpoints is fundamental from a practical perspective, because it allows global properties of interest to be monitored on subsets of nodes instead of on \( N \) as a whole. The goal of detecting an unconditional breakpoint is to find a global state \( \varphi \) such that \( \varphi [t] \) for all \( n \in N \) such that \( lub < \infty \). If no such global state exists, then the detection algorithm must be able to report this as an error.

A conditional breakpoint is specified by providing for each \( n \in N \) a local predicate, that is, a Boolean function that depends on \( n \)'s variables. The local predicate of \( n \) is denoted by \( lp_n \), and can be a constant (either \( true \) or \( false \)) if \( n \) does not participate in the breakpoint. The conditional breakpoints that we consider are either disjunctive or conjunctive. A disjunctive breakpoint is a global state at which the disjunctive predicate given by the logical disjunction of all participating nodes' local predicates is \( true \). In other words, a global state is a disjunctive breakpoint if and only if at least one of the participating nodes has a \( true \) local predicate in that global state. A conjunctive breakpoint is defined likewise, being a global state at which the conjunctive predicate given by the logical conjunction of all participating nodes' local predicates is \( true \). Put differently, a global state is a conjunctive breakpoint if and only if all participating nodes have a \( true \) local predicate in that global state.

The goal of detecting a disjunctive breakpoint is that of finding a global state \( \varphi \) such that \( lp_i = true \) at local time \( \varphi [t] \) for at least one node \( n \) that participates in the breakpoint. In the same vein, the goal of detecting a conjunctive breakpoint is to find a global state \( \varphi \) such that \( lp_i = true \) at time \( \varphi [t] \) for all nodes \( n \) that participate in the breakpoint. If node \( n \) does not participate in a breakpoint, then it suffices to set \( lp_i \) to \( false \) in the disjunctive case, or to \( true \) in the conjunctive case, for the goal of the corresponding detections to be re-stated more simply as follows. Detecting a disjunctive breakpoint is to find a global state \( \varphi \) such that \( lp_i = true \) at time \( \varphi [t] \) for at least one \( n \in N \); in the conjunctive case, it is to find a global state \( \varphi \) such that \( lp_i = true \) at time \( \varphi [t] \) for all \( n \in N \).

The following is how Sections 9.3.2 and 9.3.3 are organized. In Section 9.3.2, a trace-based algorithm is presented to halt an execution at the earliest conjunctive breakpoint. Section 9.3.3, which is where most of our study is concentrated, contains distributed algorithms for the detection of earliest disjunctive breakpoints, earliest unconditional breakpoints, and earliest conjunctive breakpoints. In the case of unconditional breakpoints, requiring the earliest such breakpoint to be detected is only meaningful if there is at least one node that does not participate in the breakpoint. As we have seen earlier in this section, nodes like this have the \( lub \) variables set to infinity, and may then be required to appear in the detected breakpoint with as early a local state as possible.

Most of the algorithms that we study employ messages other than the messages of the computation proper, and for this reason a distinction must be made between such messages and those additional messages that the algorithms employ. Messages of the computation proper are then referred to as \( comp_msg \)’s.

9.3.2 A trace-based technique

In this section, we discuss a trace-based technique to halt the execution of an asynchronous algorithm at the earliest conjunctive breakpoint that occurs. It is trace-based because, as in
Section 9.2, it is based on two phases, the first one being the recording of a trace and the second one a re-execution with special attention to the detection of the earliest conjunctive breakpoint. The trace recording is achieved precisely as in Algorithm A_Record_Trace of Section 9.2, where the variables queue and first_in_queue, respectively a queue of node references and a pointer to its first element, are employed by node \( n \) to record the neighborhood-wide order according to which it receives comp_msg's. The difference in this case is that not only a deterministic re-execution is sought, but a re-execution that does not progress beyond the earliest conjunctive breakpoint. As in Section 9.2, G's edges are assumed to be FIFO. Also, for the sake of simplicity when writing the algorithm, we assume that a node's local predicate can only become true after the node has computed and sent messages out (unless the node does not participate in the breakpoint, in which case its local predicate is perpetually true).

The approach that we adopt has the following essential ingredients. During the re-execution phase, node \( n \) may be active or inactive. It is active initially and becomes inactive when \( lp \) becomes true. A node only receives comp_msg's (and therefore only computes and sends comp_msg's out) if it is active, so that becoming inactive when its local predicate becomes true is a means of cooperating for the execution to halt at the earliest conjunctive breakpoint. The problem with such a naive way of cooperating is that, by becoming inactive and therefore not sending any comp_msg's out, a node may be precluding other nodes from reaching local states in which their local predicates can become true as well, which is an absolute must if the execution is to halt at a conjunctive breakpoint.

One way to go around this difficulty is the following. For \( n \in Neig \), node \( n \) maintains two counters, \( \text{last_received}^i_j \) and \( \text{last_sent}^i_j \), both initially set to zero, to indicate respectively the number of comp_msg's received from \( n \) and the number of comp_msg's sent to \( n \). Whenever \( n \) is finished with computing and sending comp_msg's out, either initially or in response to the reception of a comp_msg, and has remained active, it sends a request\((i+1)\) message to \( n \) such that \( n = \text{first_in_queue} \). This message is intended to activate \( n \), if it is inactive, so that it can send \( n \) the comp_msg that it needs to proceed according to the trace. When the request message reaches \( n_i \) with an \( x \) parameter, then it must be that \( x - 1 \leq \)

\[
\text{last_sent}^i_j = x - 1
\]

sent, node \( n_i \) employs the variable

\[
\text{owed}^j_i
\]

, initially equal to zero, to indicate the number of comp_msg's that need to be sent to \( n \in Neig \) before it may become inactive.

The problem that still persists with this approach is that, because edges are FIFO (as they must be for correct re-execution) and a node only receives comp_msg's when it is active and the origin of the comp_msg coincides with the node's first_in_queue variable, it may happen that a request never reaches its destination. The final fix is then to allow a node to receive all comp_msg's that reach it, and then to queue them up internally (along with their origins) on edge-specific queues if the node happens to be inactive or the comp_msg that arrived is not the one that was expected for re-execution. Upon receipt of a request from \( n \), an inactive node \( n \) works on the messages in those queues until the re-execution can no longer
progress or 
\[ \text{owed}_i^j = 0 \text{ and } l_p = \text{true}. \]
If \( i > 0 \text{ or } l_p = \text{false} \) when \( n \) exits this loop, then \( n \) becomes active and sends out a request. The reader should reflect on the reasons why this procedure ensures that the re-execution halts (i.e., all nodes become inactive) at the earliest conjunctive breakpoint (cf. Exercise 2).

Algorithm \( \text{A}_\text{Replay} \& \_\text{Halt CB} \) (*CB" for Conjunctive Breakpoint), given next, realizes the procedure we just described. In addition to the variables we already introduced, node \( n \) also employs the Boolean \( \text{active}_i \), initialized to \( \text{not } l_p \), to indicate whether it is active. Also, for each \( n_j \in \text{Neig}_i \), the queue where \( \text{comp}_\text{msg}'s \) from \( n_j \) may have to be queued is

\[
\text{msg}_\text{queue}_i^j, \text{first}_\text{in}_\text{queue}_i^j
\]
initially set to \( \text{nil} \), whose first element is assumed to be the pair ( \( \text{nil, nil} \) ), initially equal to (\( \text{nil, nil} \)), at all times. In

this algorithm, as in Algorithm \( \text{A}_\text{Replay} \), the set \( N_0 \) is given as determined by Algorithm \( \text{A}_\text{Record Trace} \).

**Algorithm \( \text{A}_\text{Replay} \& \_\text{Halt CB} \):**

**Variables:**
- \( \text{queue}_i \)
- \( \text{first}_\text{in}_\text{queue}_i \)
- \( \text{last}_\text{received}_i^j \) = 0 for all \( n_j \in \text{Neig}_i \)
- \( \text{last}_\text{sent}_i^j \) = 0 for all \( n_j \in \text{Neig}_i \)
- \( \text{owed}_i^j \) = 0 for all \( n_j \in \text{Neig}_i \)
- \( \text{active}_i = \text{not } l_p \)
- \( \text{msg}_\text{queue}_i \) = \( \text{nil} \) for all \( n_j \in \text{Neig}_i \)
- \( \text{first}_\text{origin}_\text{in}_\text{msg}_\text{queue}_i \) = \( \text{nil} \) for all \( n_j \in \text{Neig}_i \)
- \( \text{first}_\text{msg}_\text{in}_\text{msg}_\text{queue}_i \) = \( \text{nil} \) for all \( n_j \in \text{Neig}_i \)
- Other variables used by \( n_i \), and their initial values, are listed here.

**Listing 9.5**

**Input:**
- \( \text{msg} = \text{nil} \).

**Action if \( n_i \in N_0 \):**
- Do some computation;
- Send one \( \text{comp}_\text{msg} \) on each edge of a (possibly empty) subset of \( \text{Inc}_i \) and update the
  \( \text{last}_\text{sent}_i^j \)'s accordingly;
  if \( l_p \) then
  \( \text{active}_i := \text{false} \).
Listing 9.6

Input:
\[ \text{msg} = \text{comp}_\text{msg} \text{ such that } \text{origin}_\text{msg}(\text{msg}) = (n_i, n_j) \]

Action:
if \( \text{active} \) \&\& \( n_j = \text{first}_\text{in}_\text{queue} \), then
begin
\[ \text{last}_\text{received}_j^i := \text{last}_\text{received}_j^i + 1; \]
Remove \( \text{first}_\text{in}_\text{queue} \) from \( \text{queue}_i \);
Do some computation;
Send one \( \text{comp}_\text{msg} \) on each edge of a (possibly empty) subset of \( \text{Inc}_i \); and update the \( \text{owed}_k^i \)'s and the \( \text{last}_\text{sent}_k^i \)'s accordingly;
if \( \text{owed}_k^i = 0 \) for all \( n_k \in \text{Neig}_i \), then
if \( \text{lp} \), then
\( \text{active} := \text{false} \);
if \( \text{active} \), then
begin
Let \( n_k = \text{first}_\text{in}_\text{queue}_i \);
Send request\((\text{last}_\text{received}_k^i + 1)\) to \( n_k \)
end
else
Append \((n_i, \text{msg})\) to \( \text{msg}_\text{queue}_i \).
end

Listing 9.7

Input:
\[ \text{msg} = \text{request}(x) \text{ such that } \text{origin}_\text{msg}(\text{msg}) = (n_i, n_j) \]

Action:
if not \( \text{active} \) \&\& \( x - 1 = \text{last}_\text{sent}_j^i \) then
begin
\[ \text{owed}_j^i := \text{owed}_j^i + 1; \]
while ( \( \text{owed}_j^i > 0 \) \&\& \text{not} \( \text{lp} \) \&\& there exists \( n_k \in \text{Neig}_i \) such that \( \text{first}_\text{in}_\text{queue}_i \neq \text{first}_\text{origin}_\text{in}_\text{msg}_\text{queue}_i^k \))
begin
\[
\text{last\_received}^k_i := \text{last\_received}^k_i + 1;
\]
Remove first_in_queue from queue;
Remove the pair \( (\text{first\_origin\_in\_msg\_queue}^j_i, \text{first\_msg\_in\_msg\_queue}^j_i) \) from \( \text{msg\_queue}^k_i \);
Do some computation;
Send one \( \text{comp\_msg} \) on each edge of a (possibly empty) subset of \( \text{Inc}_i \) and update the
\( \text{owed}^\ell_i \)'s and the \( \text{owed}^j_i \)'s accordingly
end;
if \( \text{owed}^j_i > 0 \) or not \( \text{lp}_i \), then
begin
\text{active}_i := \text{true};
Let \( n_k = \text{first\_in\_queue}_i \)
Send \( \text{request}(\text{last\_received}^k_i + 1) \) to \( n_k \)
end
end.

There is some correspondence between the actions in this algorithm and those in Algorithm \( \text{A\_Replay} \) for simple re-execution, but they differ greatly, too. Specifically, actions (9.3) and (9.5) are related to each other, although (9.5) also undertakes the incrementing of \( \text{last\_sent}^\ell_i \) when a \( \text{comp\_msg} \) is sent to neighbor \( n_j \) and checks \( \text{lp}_i \) to see if \( n_i \) must become inactive. Likewise, actions (9.4) and (9.6) are also related to each other. The differences are in the internal queueing of \( \text{comp\_msg} \)'s and in that (9.6) increments \( \text{last\_received}^j_i \) to account for the receipt of the triggering \( \text{msg} \) on \( (n_i, n_j) \), increments \( \text{owed}^j_i \) (decrements \( \text{owed}^j_i \), if positive) upon sending a \( \text{comp\_msg} \) to neighbor \( n_k \), and in addition checks \( \text{lp}_i \) to possibly set \( \text{active}_i \) to \text{false} (if \( \text{active}_i \) remains \text{true}, then (9.6) includes the sending of a \( \text{request} \) as well). Action (9.7) deals with the reception of a \( \text{request} \) from \( n_j \). This \( \text{request} \), if indeed corresponding to a \( \text{comp\_msg} \) that was not sent, and if \( n_j \) is inactive, causes \( \text{owed}^j_i \) to be incremented and \( n_j \) to compute on its internal queues of messages. If \( \text{owed}^j_i \) remains positive or \( \text{lp}_j = \text{false} \) after this, then \( n_j \) becomes active and sends a \( \text{request} \).
As in the case of Algorithm $A_{\text{Replay}}$, the message and time complexities of Algorithm $A_{\text{Replay} \_\& \text{Halt}_{\text{CB}}}$ are the same as those of Algorithm $A_{\text{Record}_{\text{Trace}}}$. This is so because the additional request messages only increase the total number of messages exchanged and the longest causal chain of messages by a constant factor. However, the new algorithm's bit complexity may be higher, because request messages carry integers that depend on how many $\text{comp}_{\text{msg}}$'s were received by the sending node during the trace-recording phase.

We finalize the section with a couple of observations leading to issues that the reader may find worth pursuing further. The first observation is that devising a procedure similar to Algorithm $A_{\text{Replay} \_\& \text{Halt}_{\text{CB}}}$ to halt at the earliest disjunctive breakpoint during a re-execution is a very different matter. The reader is encouraged to pursue a proof that no such procedure exists, be it trace-based or otherwise (cf. Exercise 3).

As the second observation, notice that Algorithm $A_{\text{Replay} \_\& \text{Halt}_{\text{CB}}}$ does not entirely conform to the standards set by Algorithm $A_{\text{Template}}$, in the sense that in both (9.6) and (9.7) a request may follow a $\text{comp}_{\text{msg}}$ to the same node, whereas Algorithm $A_{\text{Template}}$ only allows one message to be sent to a node per action. An instructive exercise is to rewrite Algorithm $A_{\text{Replay} \_\& \text{Halt}_{\text{CB}}}$ so that this constraint is respected (cf. Exercise 4).

9.3.3 A trace-independent approach

In this section we introduce three asynchronous algorithms for the detection of breakpoints. One of the algorithms detects earliest disjunctive breakpoints, another detects earliest unconditional breakpoints, and the last one detects earliest conjunctive breakpoints whose corresponding conjunctive predicates are stable (in the sense of Chapter 6). As we remarked earlier in Section 9.3.1, not always is the requirement that an unconditional breakpoint be the earliest such breakpoint meaningful—in fact, it only makes sense when at least one node does not participate in the unconditional breakpoint. If such is not the case, then what the algorithm that we discuss achieves is the detection of the requested unconditional breakpoint.

Another pertinent observation regarding the algorithms of this section is the assumed stability of the conjunctive predicate used to detect the conjunctive breakpoints. This assumption simplifies matters tremendously, and it is in the wake of this simplicity that we adopt it. However, there do exist techniques to detect conjunctive breakpoints in the absence of stability (as in Section 9.3.2), and at least a couple of recent algorithms that do not employ traces can be found in the literature.

In contrast with the detection procedure discussed in the previous section, which was based on a trace of a previous execution, the overall approach in this section does not depend on any trace, but rather attempts to detect the required breakpoint as the computation progresses. As in the case of the global state recording discussed in Section 5.2.1, what we must handle is then the interaction of two computations on $G$. One of the computations is the computation proper, the one that progresses by the exchange of the already introduced $\text{comp}_{\text{msg}}$ messages. The other computation is the computation for breakpoint detection, which, as in Section 5.2.1, must be endowed with certain privileges with respect to the former computation, because it must be able to inspect nodes' states in that computation as well as the flow of $\text{comp}_{\text{msg}}$'s. In the case of Algorithm $A_{\text{Record}_{\text{Global}_{\text{State}}}}$, we were able to get away without being more specific about the way the two computations interacted, but in the present case we must face the need to provide some of the details, especially because the computation for breakpoint detection will have to be able to attach additional fields, in the form of parameters, to the $\text{comp}_{\text{msg}}$'s that flow among nodes.

Let us then fill in some of the details on how the two computations interact. We begin by assuming that the message complexity of the computation proper is $O(c(n, m))$, or more succinctly $O(c)$. Likewise, we assume that every node's local clock can only represent local times up to a value $T$ (i.e., $lt \leq T$ for all $n \in N$), which can be arbitrarily large but needs nevertheless be such that we can refer to it when assessing our algorithms' complexities. The following is how we henceforth assume the computation proper and the detection algorithms to interact. At $n \in N$, the actions of both computations exclude one another in time, as usual. For clarity, we view node $n$ as comprising two processes that can send messages to each other. One of the processes is referred to as $p_n$, while the other is referred to as $q_n$. For all $n \in N$, process $q_n$ is responsible for the detection algorithm, while process $p_n$ is responsible for the computation proper. Process $q_n$ may send messages to every other
process $q$, such that $n_j \in \text{Neig}$ and to $p$, while process $p$ may only send messages to $q$. In this way, $q$, is capable of intercepting every $\text{comp\_msg}$ that $p$ sends or receives, for the purpose of detecting breakpoints. In doing so, $q$, may add fields to, or strip fields off, the $\text{comp\_msg}$'s that it intercepts. Of course, every $\text{comp\_msg}$ sent by $p$, must somehow contain an indication of which $p$, it is destined to, so that $q$, can forward it appropriately to $q$. Process $q$, responds to messages it receives either on edges $(n_k, n_l) \in \mathcal{E}$ or “internally” from $p$. When specifying the algorithms for breakpoint detection, we provide actions for process $q$, only, and then mentions to the set $N_q$, never appear, as clearly actions performed by nodes in this set are actions of the computation proper.

Perhaps the most important assumption on how processes $p$, and $q$, interact is that the atomicity of $p$,’s actions may be violated, in the following sense. Process $q$, is activated (and then $p$, is suspended, while $lt$, remains constant) when $lt$, becomes equal to $lub$, (in the case of unconditional breakpoints) or when the local predicate $lp$, becomes true (in the case of conditional breakpoints), or yet upon the sending by $p$, of a $\text{comp\_msg}$ or the arrival of a message from $q$, such that $n_j \in \text{Neig}$. What this amounts to is that $q$, has some sort of "preemptive priority" over $p$.

When evaluating our algorithms’ complexities, this dual character of a node's behavior may at first seem confusing, so that it is advisable to spell out the criteria to be used right away. All the complexity measures to be given in this section are measures related to the breakpoint detection computation only, and then may involve messages sent especially for detection purposes as well as $\text{comp\_msg}$'s to which additional fields were attached (although in the latter case only the message and bit complexities are affected, not the global time complexity, which is already accounted for by the computation proper). Such measures are then aimed at capturing the “overhead” of breakpoint detection only.

Our algorithms for breakpoint detection are based on the following general approach. For $1 \leq i \leq n$, process $q$, maintains an array $gs_i$ of length $n$ representing its view of the global state to be detected. This array is initialized with zeroes (representing the earliest global state of the computation) and is updated when information is received concerning the other nodes' local unconditional breakpoints or local predicates. Such information is conveyed from node to node either by means of special broadcast messages or as additional fields attached to the $\text{comp\_msg}$'s that constitute the communication traffic of the computation proper. This information, when sent by $q$, comprises the array $gs_i$ and may, depending on the type of breakpoint to be detected, comprise additional data as well. In each of the cases we consider, this information is exchanged among nodes in such a way as to allow at least one node, say $n_x \in N$, to detect locally that the breakpoint has occurred at the global state recorded in the current $gs_x$ maintained by $q_x$, which is in all cases the earliest global state at which the breakpoint can be said to have occurred.

Let us now examine Algorithm $A_{\text{Detect\_DB}}$ ("DB" for Disjunctive Breakpoint) for the detection of disjunctive breakpoints. Such a breakpoint is a global state at which for at least one of the participating nodes the local predicate holds. Clearly, the earliest global state at which a disjunctive predicate holds does not

![Image](image-url)

**Figure 9.1** In this figure, the solid segment in a process’s horizontal line indicates the time interval during which the corresponding local predicate is true. The two cuts shown clearly correspond to global states, in fact earliest global states in which the disjunctive predicate holds.

have to be unique (Figure 9.1), as we mentioned in Section 9.3.1, so it is conceivable that more than one node detects the occurrence of the breakpoint, however at different global states.

Because of the inherent ease with which disjunctive predicates can be detected in a distributed fashion, Algorithm $A_{\text{Detect\_DB}}$ is quite straightforward. It does not employ any broadcast messages, and attaches the array $gs_i(lt)$, in addition to a "status bit" (to be discussed shortly), to the $\text{comp\_msg}$'s sent by process $q$, on behalf of $p$. This array is identical to $gs_i$ in all components except the $i$th, which is given by $lt$. Our earlier assumptions
Algorithm A_Detect_DB:

Variables:
- $gs[i][k] = 0$ for all $n_i \in N$
- $found_i = false$

Lemma 9.1.
For all $n_i \in N$, if $gs[i]$ is a global state such that $gs[i] < lt$, and no message is received at $p_i$ at time $t$ such that $gs[i] < t \leq lt$, then $gs[lt]$ is also a global state.

Proof:
If $gs[i]$ is not a global state, then there must exist $n_i, n_j \in N$ such that a $comp_msg$ was sent by $p_i$ strictly later than $gs[i][k]$ and received at $p_j$ earlier than (or at) $gs[i][f]$. By the definition of $gs[lt]$, and by hypothesis, it follows that the message must have been sent later than $gs[k]$ and arrived at $p_j$ earlier than (or at) $gs[f]$, and then $gs$, must not be a global state, which is a contradiction.

Lemma 9.2.
If $\varphi$ and $\varphi'$ are global states, then the component-wise maximum of the two is also a global state.

Proof:
Let $\varphi''$ be the component-wise maximum of $\varphi$ and $\varphi'$, and suppose that it is not a global state. Then there must exist $n_i, n_j \in N$ such that a message was sent by $p_k$ strictly later than $\varphi''[k]$ and received at $p_l$ earlier than (or at) $\varphi''[f]$. Because $\varphi''[k] \geq \varphi'[k]$ and $\varphi''[f] \geq \varphi'[f]$, then $\varphi$ must not be a global state if $\varphi''[\ell] = \varphi'[\ell]$. Likewise, if $\varphi''[\ell] = \varphi'[\ell]$, then $\varphi'$ must not be a global state.

Either case yields a contradiction.

The essence of Algorithm A_Detect_DB is the following for $n_i \in N$. Variable $lp_i$ is initialized with false at process $p_i$, and is assumed never to become true if $n_i$ does not participate in the breakpoint. Whenever $q_i$ detects that $lp_i$ has become true, it sets $gs[i]$ to $lt$ and declares the disjunctive breakpoint detected at the global state $gs$. Because every $comp_msg$ it received from $q_j$ such that $n_j \in Neig$ prior to $lt$ carried a copy of $q_j$'s view of the global state with $j$th component updated to the time the message was sent, $gs$, must indeed be a global state by Lemmas 9.1 and 9.2. In order to ensure that it is also an earliest global state with respect to the disjunctive predicate, the simple procedure we just described must only be allowed to be performed if no other node has already detected a global state that renders the one $q_i$ would detect not an earliest one. This is where the "status bit" comes in. This bit will indicate, upon arriving along with a $comp_msg$, whether any other such global state has already been detected.

Algorithm A_Detect_DB is presented next. Two additional variables employed by the algorithm are the Booleans $found$ and $found_elsewhere$, both initially set to false, which indicate respectively whether $q_i$ has detected the disjunctive breakpoint and whether such a breakpoint has already been detected elsewhere so that the one detected by $q_i$ would necessarily not be an earliest one.
\[ \text{found\_elsewhere} = \text{false}. \]

**Listing 9.8**

**Input:**
\[ msg_i = \text{nil}. \]

**Action when \( lp_i \) becomes true:**
\[
\text{if not (found}_i \text{or found\_elsewhere}_i) \text{ then}
\begin{align*}
\text{begin} & \\
& \text{gs}[i] := \text{lt}_i; \\
& \text{found}_i := \text{true} \\
\text{end.}
\end{align*}
\]

**Listing 9.9**

**Input:**
\[ msg_i = \text{comp\_msg from } p_i \text{ to } p_j. \]

**Action:**
\[ \text{Send } \text{comp\_msg(found}_i \text{or found\_elsewhere}_i, \text{gs}(\text{lt}_i)) \text{ to } q_i. \]

**Listing 9.10**

**Input:**
\[ msg_i = \text{comp\_msg(b, gs)}. \]

**Action:**
\[
\text{found\_elsewhere}_i := b \text{ or found\_elsewhere}_i; \\
\text{if not (found}_i \text{or found\_elsewhere}_i) \text{ then}
\begin{align*}
\text{for } k := 1 \text{ to } n & \text{ do}
\begin{align*}
\text{if } \text{gs}[k] < \text{gs}[k] & \text{ then}
& \text{gs}[k] := \text{gs}[k];
\end{align*}
\end{align*}
\]
\[ \text{Send } \text{comp\_msg to } p_i. \]

The next theorem establishes the correctness of Algorithm \textit{A\_Detect\_DB}. This theorem, like the others to follow in this section, state the equivalence of several conditions. The proof strategy in all theorems is then to show that the first condition implies the second, which implies the third, and so on, and finally that the last condition implies the first.

**Theorem 9.3.**
There exist \( i \in N \) and \( \geq 0 \) such that the following three conditions are equivalent to one another for Algorithm \textit{A\_Detect\_DB}.

- (i) There exists a global state \( s \) such that \( lp_i = \text{true} \) at time \( \text{lt}_i \) for at least one \( n_i \in N \).
- (ii) \( \text{found}_i \) becomes \text{true} at time \( \text{lt}_i = t \).
• (iii) At time \( t \), \( gs \) is the earliest global state at which \( lp_i = \text{true} \) for at least one \( n_i \in N \).

Proof:

(i) \( \rightarrow \) (ii):
At least one of the nodes \( n_i \) for which \( lp_i \) ever becomes true must by actions (9.9) and (9.10) have reached this state for the first time when \( \text{found\_elsewhere}_i = \text{false} \). The assertion then follows immediately by action (9.8), with \( n_i \) being this particular node and \( t \) being the local time at which \( lp_i \) becomes true for the first time.

(ii) \( \rightarrow \) (iii):
By hypothesis and by action (9.8), \( \text{found\_elsewhere} \) can only have become true after time \( t \). By Lemmas 9.1 and 9.2, the \( gs \) produced by action (9.8), the \( gs(t_l) \) used in action (9.9), and the \( gs \) yielded by action (9.10) must all be global states. As a consequence of this, by action (9.8) \( gs \) is at time \( t \) a global state at which \( lp_i = \text{true} \). If \( gs \) were not an earliest global state at which \( lp_i = \text{true} \) for at least one \( n_i \in N \), then either \( \text{found\_elsewhere} \) would by actions (9.9) and (9.10) have become true prior to \( t \), and then \( \text{found} \) would be false at \( t \); which is a contradiction, or \( lp_i \) would for some \( n_i \in N \) be true right from the start, which is ruled out by our assumption on the initial values of these variables.

(iii) \( \rightarrow \) (i):
This is immediate.

Each of the \( O(c) \) comp_msg's carries an \( n \)-component array, each of whose components is an integer no larger than \( T \), so the bit complexity of Algorithm A_Detect_DB is \( O(cn \log T) \). Because only comp_msg's are employed, the algorithm's message and global time complexity are of \( O(1) \). Each message reception requires \( O(n) \) comparisons, which is then the algorithm's local time complexity.

Detecting the other types of breakpoints we consider in this section is a considerably more intricate task in comparison with the detection of disjunctive breakpoints. These other cases comprise unconditional breakpoints and conjunctive breakpoints on stable conjunctive predicates, all of which require some sort of additional "global" information to be monitored. It is the propagation of this global information that makes use of the broadcast messages we introduced earlier.

In general, in addition to \( gs \) process \( q \) also maintains another array of Booleans with its local view of the global condition to be monitored and detected.

When disseminated by \( q \), this array is always accompanied by \( gs \) as well, so that whenever \( q \) detects locally that the global condition has occurred (by examination of its array), it also associates the contents of \( gs \) with the global state at which the condition occurred.

Messages of the broadcast type are sent by \( q \) whenever \( n \in \) is one of the nodes participating in the global condition to be detected and either its local unconditional breakpoint is reached (in the case of unconditional breakpoint detection) or its local predicate becomes true (in the case of the detection of conjunctive breakpoints). The broadcast we employ follows closely Algorithm A_PI of Section 4.1.1, but during the propagation of information an arriving \( gs \) from some process \( q \) is used by \( q \) to update \( gs \). In addition, \( gs \) and the other array accompanying it are used to update the local view at \( q \) of the global condition being monitored.

What further differentiates the broadcast that we employ in this section from Algorithm A_PI is that we adopt a "forward-when-true" rule for the propagation of information. This rule states that a process participates in the broadcast (i.e., forwards the information it receives) only when its local condition (local unconditional breakpoint reached or local predicate become true) holds. Clearly, if no comp_msg's were ever sent, then this broadcast would suffice for the detection of the desired type of breakpoint. In such a case, whichever process produced an array with true values for all the participating processes would declare the breakpoint detected at the global state given by the global-state array obtained along with it.

Algorithm A_Broadcast_When_True does this detection in the absence of comp_msg's, so long as the global condition under monitoring is stable. In this algorithm, process \( p \) maintains a Boolean variable \( lc \) to indicate whether the local condition with which \( n_i \) participates (if at all) in the global condition to be detected is true. It is initialized with false if \( n_i \) does indeed
participate in the global condition, or with \textit{true} otherwise. Stability then means that no \( n_k \in N \) exists such that \( lc_i \) is reset to \textit{false} once it becomes \textit{true}. The array associated with \( q_i \)'s view of the global condition is denoted by \( gc_i \). For \( 1 \leq k \leq n \), \( gc_i[k] \) is initialized with the same value assigned initially to \( lc_i \). Only \textit{broadcast} messages are employed in this algorithm (as the computation proper does not employ any), and are sent as \textit{broadcast}(\( gc_i \), \( gs_i \)) when \( q_i \) is the sender. As in the case of Algorithm \textit{A\_Detect\_DB} discussed earlier, a Boolean variable \textit{found}, set to \textit{false} initially, is employed to indicate whether \( q_i \) has detected the occurrence of the global condition. In addition, another Boolean variable, \textit{changed}, is used by \( q_i \) to ensure that a \textit{broadcast} message is never sent to a node if not different than the last message sent to that node.

\textbf{Algorithm A\_Broadcast\_When\_True:}

\begin{verbatim}
Variables:
gs_i[k] = 0 for all \( n_k \in N \);
gc_i[k] for all \( n_k \in N \);
found = false;
changed.

Listing 9.11

Input:
msg = nil.
Action when \( lc_i \) becomes true:
gc_i := lc_i;
gs_i := lt;
if gc_i[1] \& \ldots \& gc_i[n] then
  found := true
else
  Send \textit{broadcast} (gc_i, gs_i) to all \( q_j \) such that \( n_j \in \text{Neig}_i \).

Listing 9.12

Input:
msg = (gc, gs).
Action:
if not found, then
  begin
    changed := false;
    for \( k := 1 \) to \( n \) do
      if gs[k] < gs[k] then
        begin
          gs[k] := gs[k];
gc[k] := gc[k];
        changed := true
        end;
    if \( lc_i \) and changed, then
      if gc_i[1] \& \ldots \& gc_i[n] then
        found := true
      else
        Send \textit{broadcast} (gc_i, gs_i) to all \( q_j \) such that \( n_j \in \text{Neig}_i \).
  end;
\end{verbatim}
Properties of Algorithm \textit{A\_Broadcast\_When\_True} are given in the following theorem.

**Theorem 9.4.**

There exist $n_i \in N$ and $t \geq 0$ such that the following three conditions are equivalent to one another for Algorithm \textit{A\_Broadcast\_When\_True}.

1. There exists a global state $\mathcal{G}$ such that $l_{c_k} = \text{true}$ at time $\mathcal{G}[k]$ for all $n_k \in N$.
2. $\text{found}_i$ becomes $\text{true}$ at time $l_t = t$.
3. At time $l_t = t$, $g_s$ is the earliest global state at which $l_{c_k} = \text{true}$ for all $n_k \in N$.

**Proof:**

(i) $\Rightarrow$ (ii):

If exactly one node participates in the global condition, then by action (9.11) $\text{found}_i$, becomes $\text{true}$, with $n_i \in N$ being this node and $t$ the time at which $l_{c_i}$ becomes $\text{true}$. No messages are ever sent in this case. If at least two nodes participate, then at least one of them, say $n_k \in N$, is such that $q_k$ does by action (9.11) send a broadcast message to $n_k$'s neighbors when $l_{c_k}$ becomes $\text{true}$, which by action (9.12) pass the updated information on, so long as the update introduced changes and their local conditions hold as well. Because this broadcast carries $l_{c_k}$, it must introduce changes when reaching every node for the first time and is therefore propagated. This happens to the local condition of every participating node, and then at least one process, say $q_i$, upon having been reached by their broadcasts, and having $l_{c_i} = \text{true}$, sets $\text{found}_i = \text{true}$. The value of $t$ here is either the time at which the last broadcast to reach $q_i$ does reach it by action (9.12) or the time at which $l_{c_i}$ becomes $\text{true}$ by action (9.11).

(ii) $\Rightarrow$ (iii):

By Lemmas 9.1 and 9.2, the $g_s$ produced in actions (9.11) and (9.12) are global states. Consequently, and by actions (9.11) and (9.12) as well, at time $t$ $g_s$ is a global state at which $l_{c_i} = \text{true}$ for all $n_i \in N$. That $g_s$ is the earliest such global state is immediate, because of the absence of $\text{comp\_msg}'s$, which implies that $g_s[k]$ is either zero or the time at which $l_{c_k}$ becomes $\text{true}$.

(iii) $\Rightarrow$ (i):

This is immediate.

Let us now assess Algorithm \textit{A\_Broadcast\_When\_True}'s complexities. The worst case is that in which all nodes start the algorithm concurrently, and furthermore the broadcast started by a node traverses all edges. The algorithm's message complexity is then $O(nm)$. Because two $n$-component arrays are sent along with each message, one comprising single-bit components, the other integers bounded by $T$, the bit complexity becomes $O(nm \log T)$. No causal chain of messages comprises more than $O(n)$ messages, because this is what it takes for a broadcast to reach all nodes, so this is the algorithm's global time complexity. The local time complexity is like that of Algorithm \textit{A\_Detect\_DB}, therefore of $O(n)$.

Algorithms \textit{A\_Detect\_DB} and \textit{A\_Broadcast\_When\_True} detect breakpoints in two extreme situations, respectively when the breakpoint is a disjunctive breakpoint and when the breakpoint is a conjunctive breakpoint but the computation proper does not ever send any message (it is simple to note that the case of unconditional breakpoints in the absence of $\text{comp\_msg}'s$ is in fact a case of conjunctive breakpoints). In the former case only are $\text{comp\_msg}'s$ employed, whereas in the latter case only broadcast messages are needed. Other situations between these two extremes are examined in the sequel, and then the messages received by process $q_i$ are either like $\text{comp\_msg}(g_c, g_s)$ or like broadcast$(g_c, g_s)$. 
Now we introduce \textit{A\_Detect\_UB} ("UB" for Unconditional Breakpoint), a distributed algorithm to detect the occurrence in a distributed computation of an unconditional breakpoint. As we discussed previously, this unconditional breakpoint is specified, for each node actually participating in the breakpoint, as a local time denoted by \textit{lub}, for \( n_i \in N \). For nodes \( n_i \) that do not participate in the breakpoint, we have chosen to adopt \( lub_i = \infty \), so that \( lt_i \) can never equal \( lub_i \).

Algorithm \textit{A\_Detect\_UB} must operate somewhere between the two extreme situations assumed by Algorithms \textit{A\_Detect\_DB} and \textit{A\_Broadcast\_When\_True}, and then can be regarded as a mixture of those. Put differently, the detection of unconditional breakpoints does require the detection of a global condition (which is ruled out by Algorithm \textit{A\_Detect\_DB}) and must be applicable to the case when messages of the computation proper exist (which are disallowed by Algorithm \textit{A\_Broadcast\_When\_True}).

The variables employed by Algorithm \textit{A\_Detect\_UB} are essentially the ones introduced earlier for the other two algorithms, except that for process \( p_i \), the Boolean variable \( ic \) is now replaced with the occurrence of the equality \( lt_i = lub_i \), and furthermore the array \( ub \), used to indicate \( q \)'s view of the occurrence of the local unconditional breakpoints at all nodes, is now used in lieu of the array \( gc \). For \( n_i \in N, ub[k_i] \) may be either \textit{true}, \textit{false}, or \textit{undefined}. It is \textit{true} or \textit{false} if \( n_i \) participates in the unconditional breakpoint and is viewed at \( n \)'s having already reached its local unconditional breakpoint or not, respectively, and is \textit{undefined} if \( n_k \) is not one of the nodes participating in the unconditional breakpoint. Initially, \( ub[k_i] \) is set to \textit{false} for every participating \( n_i \), and to \textit{undefined} if \( n_i \) does not participate.

Algorithm \textit{A\_Detect\_UB} proceeds as follows. Whenever \( q_i \) detects that \( lt_i = lub_i \), it updates \( ub_i \) \[i\] and gs[i] accordingly and starts a broadcast to disseminate the updated \( ub \) and \( gs \). This broadcast proceeds like the one in Algorithm \textit{A\_Broadcast\_When\_True}, i.e., it is never forwarded by a node whose local unconditional breakpoint has not yet been reached (unless the node does not participate in the unconditional breakpoint), and in addition no duplicate information is ever forwarded by any node. Every \textit{comp\_msg} is sent with \( ub \) and \( gs(lt) \) attached to them, in the way of Algorithm \textit{A\_Detect\_DB}, so that the global state that is eventually detected is indeed a global state. This detection, if achieved by \( q_i \), corresponds to the verification that \( ub[k_i] \neq \textit{false} \) for all \( n_i \in N \), that is, every node has either reached its local unconditional breakpoint or is not participating in the unconditional breakpoint.

One of the difficulties in designing Algorithm \textit{A\_Detect\_UB} is that it must be able to detect situations in which the requested set of local unconditional breakpoints does not constitute a global state (Figure 9.2). In such situations, an error must be reported and the computation proper must be allowed to progress normally. The detection of such a situation can be achieved along the following lines. Suppose \( q_i \) receives a \textit{comp\_msg}(\( ub, gs \)) from some process \( q_j \). If \( ub[j] = \textit{true} \) and \( ub[j] = \textit{false} \) at this moment, then clearly an error has occurred in the determination of the unconditional breakpoint, as \( p_i \) will never reach its local unconditional breakpoint in such a way that is consistent with the local unconditional breakpoint of \( p_j \) from the point of view of a global state.

The possibility of having nodes for which no local unconditional breakpoint is specified complicates the treatment of these erroneous conditions a little bit. If a causal chain of \textit{comp\_msg}'s beginning at \( q_i \) such that \( ub[t] = \textit{true} \) and going through a number of processes \( q_j \), for which \( ub[k] = \textit{undefined} \) eventually leads to \( q_j \) such that \( ub[j] = \textit{false} \), then an error must be detected just as in the case discussed earlier. The way we approach this is by artificially setting \( ub[k] \) to \textit{true} for all the \( q_j \)'s. A Boolean variable \( in\_error \), initially set to \textit{false}, is employed by \( q_i \) to indicate whether an erroneous condition has been detected.

Nodes that do not participate in the unconditional breakpoint also complicate the detection of earliest global states. If such nodes did not exist, or if we did not require the earliest global state to be detected when they did exist, then what we have outlined so far would suffice for Algorithm \textit{A\_Detect\_UB} to work as needed.

\begin{figure}[h]
\centering
\includegraphics[width=0.5\textwidth]{figure92.png}
\caption{The tiny solid segment in a process's horizontal line indicates the local time to which...}
\end{figure}
the corresponding local unconditional breakpoint has been set. Clearly, the settings in this figure are erroneous, as the cut (shown as a dashed line) that goes through them does not correspond to a global state.

However, the existence of causal chains of $\text{comp}_\text{msg}$'s similar to the one we just described but beginning at $q_\ell$ such that $\text{ub}[\ell] = \text{undefined}$ may lead to distinct earliest global states, depending on whether it leads to $q_\ell$, such that $\text{ub}[\ell] = \text{false}$ or $\text{ub}[\ell] = \text{true}$ (Figure 9.3). Only in the former case should $q_\ell$ take into account what it receives attached to the $\text{comp}_\text{msg}$ in updating $g_s$, but the senders of the preceding messages in the causal chain have no way of knowing this beforehand. The strategy we adopt to tackle this is the following. In addition to maintaining $g_s$ as a local view of the global state to be detected, $q_\ell$ also maintains an alternative view, denoted by $\text{alt}_\text{gs}_i$, which is initialized like $g_s$, but only updated or attached to outgoing $\text{comp}_\text{msg}$'s (the latter in place of $g_s$) if $\text{ub}[\ell] = \text{undefined}$. Arriving $\text{comp}_\text{msg}$'s at $q_\ell$ affect $g_s$, if $\text{ub}[\ell] = \text{false}$ or $\text{alt}_\text{gs}$, if $\text{ub}[\ell] = \text{undefined}$. So for $q_\ell$ such that $\text{ub}[\ell] = \text{undefined}$, $g_s[k] \leq \text{alt}_\text{gs}[k]$ for all $n_\ell \in N$, and therefore $g_s$, may constitute an earlier global state than $\text{alt}_\text{gs}$.

![Figure 9.3](image)

**Figure 9.3:** Following the same conventions as in Figure 9.2, here a situation is depicted in which only one node participates in the unconditional breakpoint (node $n_\ell$). Depending on how the corresponding local unconditional breakpoint is placed with respect to the reception of the message by $p_\ell$, the other processes appear in the resulting earliest global state differently, as shown in parts (a) and (b).

**Algorithm A_Detect_UB:**

**Variables:**
- $g_s[k] = 0$ for all $n_\ell \in N$;
- $\text{ub}[\ell] = \text{false}$;
- $\text{found}_i = \text{false}$;
- $\text{changed}_i$;
- $\text{in\_error}_i = \text{false}$;
- $\text{alt}_\text{gs}[k] = 0$ for all $n_\ell \in N$.

**Listing 9.13**

**Input:**
msg\_i = \text{nil}.

**Action when detecting that \(lt_i = lub\):**

\[
\text{if not } in\_\text{error, then}
\begin{align*}
\ & ub[i] := \text{true}; \\
\ & gs[i] := lt_i; \\
\ & \text{if } ub[k] \neq \text{false for all } k = 1,\ldots,n \text{ then} \\
\ & \quad \text{found} := \text{true} \\
\ & \text{else} \\
\ & \quad \text{Send } broadcast(ub_i, gs_i) \text{ to all } q_j \text{ such that } n_j, \in Neig_i
\end{align*}
\text{end.}
\]

**Listing 9.14**

\[
\begin{align*}
\text{Input:} & \quad \text{msg}_i = broadcast(ub, gs). \\
\text{Action:} & \quad \text{if not } (in\_\text{error, or } found) \text{ then} \\
\ & \begin{align*}
\ & \text{changed} := \text{false}; \\
\ & \text{for } k := 1 \text{ to } n \text{ do} \\
\ & \quad \text{if } gs[k] < gs[k] \text{ then} \\
\ & \quad \quad \begin{align*}
\ & \quad gs[k] := gs[k]; \\
\ & \quad ub[k] := ub[k]; \\
\ & \quad \text{changed} := \text{true}
\end{align*} \\
\ & \text{if } ub[i] = \text{undefined then} \\
\ & \text{for } k := 1 \text{ to } n \text{ do} \\
\ & \quad \text{if } alt\text{gs}[k] < gs[k] \text{ then} \\
\ & \quad \quad alt\text{gs}[k] := gs[k]; \\
\ & \text{if } lub_i \neq \text{false and } changed \text{ then} \\
\ & \text{if } ub[k] \neq \text{false for all } k = 1,\ldots, n \text{ then} \\
\ & \quad \text{found} := \text{true} \\
\ & \text{else} \\
\ & \quad \text{Send } broadcast(ub_i, gs_i) \text{ to all } q_j \text{ such that } n_j, \in Neig_i
\end{align*}
\end{align*}
\text{end.}
\]

**Listing 9.15**

\[
\begin{align*}
\text{Input:} & \quad \text{msg}_i = \text{comp\_msg} \text{ from } p_i \text{ to } p_j. \\
\text{Action:} & \quad \text{if } ub[i] = \text{undefined then} \\
\ & \text{Send } \text{comp\_msg}(ub, alt\text{gs}(lt_i)) \text{ to } q_i \\
\ & \text{else} \\
\ & \text{Send } \text{comp\_msg}(ub, gs(lt_i)) \text{ to } q_i
\end{align*}
\]
Next we give properties of Algorithm A_Detect_UB related to its correctness.

**Theorem 9.5.**

There exist \( n_k \in N \) and \( t \geq 0 \) such that the following four conditions are equivalent to one another for Algorithm A_Detect_UB.

1. (i) There exists a global state \( \mathcal{G} \) such that \( G[k] = lub_k \) for every \( n_k \in N \) such that \( lub_k < \infty \).
2. (ii) in_error never becomes true for any \( n_k \in N \).
3. (iii) found, becomes true at time \( lt = t \).
4. (iv) At time \( lt = t \), \( gs \) is the earliest global state at which \( gs[k] = lub_k \) for every \( n_k \in N \) such that \( lub_k < \infty \).

**Proof:**

(i) \( \iff \) (ii):

Suppose that there does exist \( n_k \in N \) such that in_error becomes true. By action (9.16), this must happen upon receipt, when \( ub[k] = \text{false} \), of a comp_msg contained in a causal chain of comp_msg's started at, say, process \( q_t \), sent when \( ub[t] = \text{true} \). No array \( \mathcal{G} \) such that \( G[k] = lub_k \) and \( G[t] = lub_t \) can then be a global state, and because both \( lub_k < \infty \) and \( lub_t < \infty \), we have a contradiction.

(ii) \( \iff \) (iii):

If in_error never becomes true for any \( n_k \in N \), then actions (9.13) and (9.14) are, so far as broadcast messages are concerned, identical to actions (9.11) and (9.12), respectively, of Algorithm A_Broadcast_When_True. This part of the proof is then analogous to the (i) \( \iff \) (ii) part in the proof of Theorem 9.4.
By Lemmas 9.1 and 9.2, the gs produced by action (9.13), the gs(lt) and alt_gs(lt) used in action (9.15), and the gs, and alt_gs, produced by actions (9.14) and (9.16) are all global states. This implies, by actions (9.13) and (9.14) and at time t, that gs is a global state at which ub[k] \neq false for all n ∈ N, or, equivalently, a global state such that gs[k] = lub for every n ∈ N such that lub < ∞. In order to show that gs is the earliest global state with these characteristics, consider any other n-component array of local times, call it \( \varphi \) such that

\[
\varphi[k] = gs[k] \text{ for all } n \in N \text{ such that } lub < \infty, \quad \text{and} \quad \varphi[k] < gs[k] \text{ for at least one } n \in N \text{ such that } lub_\varphi = \infty.
\]

For this particular n, in order for gs[k] to have been assigned the value greater than \( \varphi(k) \), a causal chain of comp_msg's must have existed from q, (leaving at time gs[k]) to some q, ∈ N, where by action (9.16) it must have arrived at q, when ub[lt] = false (otherwise gs, would not have been updated, and so neither would gs through the broadcast). In addition, because in_error, must have remained false, every process involved in this chain (except for q, but including q) must have had an undefined in its local record of its local unconditional breakpoint (for q, ub[k] = undefined). But because ub[lt] was found to be false, \( \varphi \) cannot possibly be a global state such that \( \varphi k = lub \) for all n ∈ N such that lub < ∞.

(iv) ⇒ (i):

This is immediate.

The message complexity of Algorithm A_Detect_UB is the same as that of Algorithm A_Broadcast_When_True, that is, O(nm). The algorithm's bit complexity is the sum of those of Algorithm A_Detect_DB and Algorithm A_Broadcast_When_True, therefore equal to O((c + nm)n log T). The global and local time complexities of Algorithm A_Detect_UB are the same as Algorithm A_Broadcast_When_True's, that is, O(n).

We now finally come to Algorithm A_Detect_CB_Stable for the detection of conjunctive breakpoints on stable conjunctive predicates. Such predicates are specified for each participating node n ∈ N as the local predicate lp, endowed with the property that it remains true once it becomes true. Unconditional breakpoints are also breakpoints on stable conjunctive predicates, but much more rigid than the ones we consider now, as in that case the detected global state is required to match the local unconditional breakpoints specified for the participating nodes exactly. In contrast, the ones we are now beginning to consider only ask that the local predicates of the participating nodes be true in the detected global state, although in some nodes they may have become true earlier than the local times given by the global state. Not surprisingly, then, the algorithm that we introduce next can be regarded as a slight simplification of Algorithm A_Detect_UB, as error conditions no longer need to be addressed.

Algorithm A_Detect_CB_Stable is in many senses related to Algorithm A_Detect_UB, and as such can also be viewed as a conceptual mixture of the principles employed in Algorithms A_Detect_DB and A_Broadcast_When_True. With respect to the latter, the local condition for n ∈ N, lc, is now expressed by the very local predicate lp, we have been considering throughout, and q's view of the global condition, gc, is now the array cp. For all n ∈ N, cp[k] is initialized like lp, that is, to false if n is participating in the breakpoint, and to true otherwise. All the other variables employed by Algorithm A_Detect_CB_Stable have the same meaning they had when used in previous contexts.

The simplification of Algorithm A_Detect_UB to yield A_Detect_CB_Stable does not go any further than the elimination of error detection, as an alternative local view at q of the global state to be detected, alt gs, is still needed to aid in the detection of the earliest global state of interest. Similarly to the case of unconditional breakpoints, a causal chain of comp_msg's beginning at q such that cp[lt] = true, going through a number of q's each with cp[k] = true as well, and finally reaching q with cp[lt] = false requires q to take into account what it receives attached to the comp_msg in updating gs. On the other hand, if no such q is ever reached, then the detected global state has a chance to be an earlier one (Figure 9.4).
Maintaining alt_gs has the function of allowing this earlier global state to be saved in gs, to be used in case no causal chain of the sort we just described ever occurs. The array alt_gs is initialized like gs and is attached to comp_msg’s with its ith

![Figure 9.4](image)

**Figure 9.4:** The conventions employed in this figure are the same as those of Figure 9.1, and the situation depicted is quite akin to that of Figure 9.3. Specifically, the earliest global state at which the conjunctive predicate holds depends on when n’s local predicate becomes true with respect to the reception of the message by p, as shown in parts (a) and (b).

A comp_msg arriving at q affects alt_gs, and may eventually affect gs, which happens if cp[i] = false upon arrival of the comp_msg, by simply updating gs to alt_gs, when lp becomes true. Only in this situation, or upon the receipt of broadcast messages, does gs get updated, but then so does alt_gs, so gs[k] ≤ alt_gs[k] for every n ∈ N.

**Algorithm A_Detect_CB_Stable:**

Variables:
- gs[k] = 0 for all n_k ∈ N;
- cp[k] for all n_k ∈ N;
- found_i = false;
- changed;
- alt_gs[k] = 0 for all n_k ∈ N.

Listing 9.17
Input:
\[ \text{msg} = \text{nil}. \]

**Action when \( lp \) becomes true:**
\[
\begin{align*}
\text{cp}[i] &:= lp; \\
\text{alt_gs}[i] &:= lt; \\
\text{for } k &:= 1 \text{ to } n \text{ do} \\
\text{gs}[k] &:= \text{alt_gs}[k]; \\
\text{if } \text{cp}[1] \land \ldots \land \text{cp}[n] \text{ then} \\
\text{found} &:= \text{true} \\
\text{else} \\
\text{Send broadcast}(cp, gs) \text{ to all } q_j \text{ such that } n_j \in \text{Neig}_i.
\end{align*}
\]

Listing 9.18

Input:
\[ \text{msg} = \text{broadcast}(cp, gs). \]

**Action:**
\[
\begin{align*}
\text{if not found, then} \\
\text{begin} \\
\text{changed} &:= \text{false}; \\
\text{for } k &:= 1 \text{ to } n \text{ do} \\
\text{if } \text{gs}[k] < \text{gs}[k] \text{ then} \\
\text{begin} \\
\text{gs}[k] &:= \text{gs}[k]; \\
\text{cp}[k] &:= \text{cp}[k]; \\
\text{changed} &:= \text{true} \\
\text{end}; \\
\text{for } k &:= 1 \text{ to } n \text{ do} \\
\text{if } \text{alt_gs}[k] < \text{gs}[k] \text{ then} \\
\text{alt_gs}[k] &:= \text{gs}[k]; \\
\text{if } \text{cp}[i] \text{ and changed, then} \\
\text{if } \text{cp}[1] \land \ldots \land \text{cp}[n] \text{ then} \\
\text{found} &:= \text{true} \\
\text{else} \\
\text{Send broadcast}(cp, gs) \text{ to all } q_j \text{ such that } n_j \in \text{Neig}_i.
\end{align*}
\]

Listing 9.19

Input:
\[ \text{msg} = \text{comp_msg} \text{ from } p_i \text{ to } p_j. \]

**Action:**
\[ \text{Send comp_msg}(cp, \text{alt_gs}(lt)) \text{ to } q_i. \]

Listing 9.20
Input:
\[ \text{msg} = \text{comp\_msg}(\text{cp}, \text{gs}). \]

Action:
\[
\begin{align*}
\text{if not found, then} \\
\text{for } k := 1 \text{ to } n \text{ do} \\
\text{if } \text{alt\_gs}[k] < \text{gs}[k] \text{ then} \\
\text{begin} \\
\text{cp}[k] := \text{cp}[k]; \\
\text{alt\_gs}[k] := \text{gs}[k] \\
\text{end}; \\
\text{Send comp\_msg to } p_i.
\end{align*}
\]

Correctness properties of Algorithm A_Detect_CB_Stable are established in the following theorem.

**Theorem 9.6.**

There exist \( n_i \in \mathbb{N} \) and \( t \leq 0 \) such that the following three conditions are equivalent to one another for Algorithm A_Detect_CB_Stable.

- (i) There exists a global state \( \varphi \) such that \( \text{lp}_k = \text{true} \) at time \( \varphi[k] \) for all \( n_k \in \mathbb{N} \).
- (ii) \( \text{found}_i \) becomes \( \text{true} \) at time \( \text{lt}_i = t \).
- (iii) At time \( \text{lt}_i = t \), \( \text{gs}_i \) is the earliest global state at which \( \text{lp}_k = \text{true} \) for all \( n_k \in \mathbb{N} \).

Proof:

(i) \( \Rightarrow \) (ii):

Actions (9.17) and (9.18) are, from the standpoint of broadcast messages alone, identical to actions (9.11) and (9.12), respectively, of Algorithm A_Broadcast_When_True. This part of the proof then goes along the same lines as the (i) \( \Rightarrow \) (ii) part in the proof of Theorem 9.4, so long as no \text{comp\_msg} overruns any broadcast message on any edge. When this happens, however, the propagation of the broadcast message may by action (9.18) be interrupted after traversing the edge, specifically upon arriving, say at process \( q_k \), and by action (9.18) finding \( \text{cp}_k[k] = \text{true} \) without causing changes to \( \text{gs}_k \) or to \( \text{cp}_k \). This is so because the \( \text{gs}_k \) carried by the broadcast message is no greater than \( \text{gs}_k \) in any component, which in turn was updated by action (9.17) when \( \text{lp}_k \) became \( \text{true} \) with the \( \text{alt\_gs}_k \) produced by action (9.20) upon receipt of the \text{comp\_msg}. The broadcast that by action (9.17) \( q_k \) then initiates when \( \text{lp}_k \) becomes \( \text{true} \) allows the proof to proceed like that of the (i) \( \Rightarrow \) (ii) part in the proof of Theorem 9.4 as well.

(ii) \( \Rightarrow \) (iii):

By Lemmas 9.1 and 9.2, the \( \text{gs}_i \) and \( \text{alt\_gs}_i \) produced by actions (9.17) and (9.18), the \( \text{alt\_gs}_i \) (lt) used in action (9.19), and the \( \text{alt\_gs}_i \) produced by action (9.20) must all be global states. A consequence of this is that, by actions (9.17) and (9.18), \( \text{gs}_i \) is at time \( t \) a global state at which \( \text{cp}_[k] = \text{true} \) for all \( n_k \in \mathbb{N} \). To show that \( \text{gs}_i \) is the earliest such global state requires that we consider any other \( n \)-component array of local times, call it \( \varphi \), such that \( \text{lp}_k = \text{true} \) at time \( \varphi[k] \) for all \( n_k \in \mathbb{N} \) and such that \( \varphi[k] < \text{gs}[k] \) for at least one \( n_k \in \mathbb{N} \). For this particular \( n \), \( \text{gs}[k] \) can only have been assigned the value greater than \( \varphi[k] \) if a causal chain of \text{comp\_msg}'s existed from \( q_k \) (leaving at time \( \text{gs}[k] \)) to some \( q_{\ell} \in \mathbb{N} \), which by action (9.17) must have arrived at \( q_{\ell} \) when \( \text{cp}_[\ell] = \text{false} \) (otherwise \( \text{gs}_{\ell} \) would not have been updated, and...
so neither would $gs$, by means of the broadcast). But because $cp[t]$ was found to be $false$, cannot possibly be a global state such that $lp_x = true$ at time $\forall t$, for all $x, \in N$.

(iii) $\Rightarrow$ (i):

This is immediate.

All the complexities of Algorithm $A_Detect_CB_Stable$ are the same as the corresponding complexities of Algorithm $A_Detect UB$. In finalizing this section, we suggest that the reader investigate simplifications to the algorithms of this section (except Algorithm $A_Broadcast_When_True$) if they are not required to detect earliest global states, but instead any global state in which the desired properties hold (cf. Exercise 5).

9.4 Exercises

1. Devise a solution for the problem discussed in Section 9.2 if the edges are not FIFO.
2. Prove that Algorithm $A_Replay_&_Halt_CP$ halts at the earliest conjunctive breakpoint.
3. Prove that there does not exist an algorithm for halting at an earliest disjunctive breakpoint, unless it is acceptable to progress further than that global state and then return by means of a rollback.
4. Rewrite Algorithm $A_Replay_&_Halt_CP$ so that no node sends more than one message to the same neighbor per action.
5. Show how to simplify the algorithms of Section 9.3.3 (except Algorithm $A_Broadcast_When_True$) given that earliest global states do not have to be detected.

1. Devise a solution for the problem discussed in Section 9.2 if the edges are not FIFO.
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5. Show how to simplify the algorithms of Section 9.3.3 (except Algorithm $A_Broadcast_When_True$) given that earliest global states do not have to be detected.

9.5 Bibliographic notes


The material in Section 9.2 is inspired in the work by LeBlanc and Mellor-Crummey (1987). Additional sources on the same problem include Netzer and Xu (1993). The sources for reference within the topic of breakpoint detection and related issues are quite abundant. The material discussed in Section 9.3.1, for example, can be enlarged by checking Cooper and Marzullo (1991) for a strong-sense approach, while the sources for additional insights into the use of arrays of local times as global states are Mattern (1989) and Fidge (1991). Publications of interest on checkpointing and rollback recovery include Kim, You, and Abouelnaga (1986), Koo and Toueg (1987), Bhargava and Lian (1988), Goldberg, Gopal, Lowry, and Strom (1991), Ramanathan and Shin (1993), and Xu and Netzer (1993).

The source of the material discussed in Section 9.3.2 is Manabe and Imase (1992), where the concern for detecting earliest global states seems to have first appeared. Section 9.3.3 is
closely based on Drummond and Barbosa (1994), where an algorithm to detect earliest conjunctive breakpoints on predicates that do not need to be stable is also presented. References for material closely related to that of Section 9.3.3 are Miller and Choi (1988a) and Garg and Waldecker (1994). Other publications on the detection of breakpoints include the one by Spezialetti (1991), which is based on an earlier formulation that appears to be prone to missing many possible global states (Spezialetti and Kearns, 1989).

Chapter 10: Simulation

Overview
In this chapter, a physical system is to be understood as a collection of physical entities, called physical processes, whose states we wish to determine for all times in a certain range, beginning at certain initial conditions. Physical processes interact with one another, and it is as a result of this interaction that their states change. Our interest is in the study of distributed algorithms for the determination by simulation of all physical processes’ states in the appropriate time range. We divide our presentation according to the nature of the state changes a physical process may undergo, after detailing the physical-system model and related notions in Section 10.1.

If the state of a physical process may change as a consequence of the processes’ continual interaction with one another (often at all time units), then the essential drive of the simulation is time itself, and the simulation is referred to as being time-stepped. Algorithms for this type of simulation are presented in Section 10.2 for two general classes of physical systems. If, on the other hand, only at some special instants do the physical processes interact by means of the so-called events that may cause the state of a physical process to change, then it is the events that drive the simulation, which is then said to be event-driven. Algorithms for the event-driven simulation of physical systems are discussed in Section 10.3 and Section 10.4 from two different perspectives on how to guarantee that the causality that exists among such events is preserved. These sections present, respectively, typical approaches of the so-called conservative and optimistic types. Section 10.5 and Section 10.6 contain brief descriptions of extensions on some of the material discussed in the earlier sections. In the case of Section 10.5, the extension is on methods for simulating systems whose timing nature is not so well defined as those treated earlier, and therefore require simulation methods that provide a hybrid approach between the time-stepped and the event-driven. In Section 10.6, we digress on a general framework encompassing all sorts of physical systems and simulation strategies discussed along the chapter.

Exercises and bibliographic notes appear, respectively, in Sections 10.7 and 10.8.

The complexities of all the algorithms to be discussed in this chapter are highly dependent upon the particular physical system being simulated. For this reason, and unlike our practice so far in the book, in this chapter we do not touch the issue of algorithm complexity at all. As in earlier occasions (cf. Section 9.3.2), it does occasionally happen during this chapter that a node sends more than one message to the same node in the same action. While this is not in full conformity with Algorithm A_Template, fixing it if necessary should be a simple matter.

10.1 Physical and logical processes

The physical system is represented by either the undirected graph 

\[ G = (N, E) \]

or the directed graph 

\[ G = (N, D) \]

where \( N \) is the set of physical processes. 

\( G \) is an undirected graph if and only if, for all \( v_i, v_j \in N \) such that \( v_i \neq v_j \) and \( v_i \) affects the state evolution of \( v_j \), it is also the case that \( v_j \) affects the state evolution of \( v_i \). In this case, \( (v_i, v_j) \in E \). Otherwise, the representation as a directed graph is chosen, and \( (v_i \rightarrow v_j) \in D \) if and only if \( v_i \) affects the state evolution of \( v_j \).
The state of a physical process \( \nu \) at time \( t \geq 0 \) is in this chapter denoted by \( x_i(t) \). The goal of a simulation by computer of the physical system is to determine \( x_i(t) \) for all \( \nu \in N \) and all time \( t \) such that \( 0 \leq t \leq T \). The value of \( T \) may be known a priori or it may have to be determined as the simulation progresses (for example, in cases in which the system is to be simulated until some sort of convergence is detected).

Physical systems are models of natural systems of interest in various scientific disciplines, so the particular natural system at hand, or the simplifications made when constructing the model, ultimately dictates the nature of the time \( t \) that governs the evolution of the physical system. Often \( t \) is continuous time (as in all cases in which differential equations are employed to build the model), but equally as often (for example, in the case of cellular automata and some other automaton networks) it is discrete. In either case, however, the simulation by computer of the physical system must be restricted to determining the states of physical processes at discrete instants (determined by the achievable precision within the particular computer system in use), and then it is legitimate to assume, for all purposes in this chapter, that \( t \) is a nonnegative integer between 0 and \( T \) (also assumed to be an integer).

The simulation of the physical system \( G \) is achieved by a logical system, which contains a logical process for each of \( G \)’s physical processes. Logical processes attempt to mimic the interaction that occurs among physical processes in the physical system, and for all times in the appropriate range output the states of the corresponding physical processes. The approach of a logical process to the simulation of the corresponding physical process depends largely on how physical processes interact. In this respect, a subdivision into two broad classes is normally employed.

If physical processes interact continually in such a way that the instants at which the state of a physical process may change can be determined beforehand, then the main drive of the logical system is time itself, and the simulation that the logical system carries out is called time-stepped. If, on the other hand, state changes in the physical processes are restricted to special instants in which they interact, and such instants can only be known as the simulation progresses, then the simulation is referred to as being event-driven. In allusion to the denomination as an event commonly employed to designate the interactions among physical processes. As the reader will have the opportunity to verify later in this section, such events do bear resemblance to the events employed in Section 3.1 to model distributed computations, but they are not the same at all.

The essence of a time-stepped simulation is quite simple. Basically, at each of the foreseeable instants at which a physical process \( \nu \)’s local state may change, the corresponding logical process does the update based on the current state of \( \nu \), and on the states of the other physical processes that exert influence on the state of \( \nu \) (these are either all the processes connected to \( \nu \) in \( G \), if \( G \) is undirected, or physical processes \( \nu \) such that \( (\nu_i \rightarrow \nu) \in D \), otherwise). Timestepped simulations are treated in Section 10.2.

The essentials of an event-driven simulation are also simple, but are best described if we resort to a sequential version of it first. A sequential event-driven simulation employs a queue of events. In this queue, events are kept in nondecreasing order of the time at which they must happen. The task of a sequential simulator is to iterate on this queue until the queue is empty or no event in it is scheduled to happen at a time no greater than \( T \). At each iteration, the simulator removes the first event from the queue, updates the state of the corresponding physical process, and possibly inserts in the queue new events to happen farther in time at some of the other physical processes. Initially, the queue contains the only events that happen spontaneously, which we assume to happen at time zero. The key ingredient in guaranteeing the correct simulation of the physical system is that events are processed in nondecreasing order of time.

In the distributed case, one possibility is for each logical process to maintain a similar queue for the events that are scheduled to happen at the physical process to which it corresponds. Initially, only a logical process at whose physical process an event happens spontaneously at time zero has a nonempty queue. The set of such physical processes is \( N_0 \subseteq N \). This
queue is maintained in increasing order of time (the reason for increasing in place of nondecreasing is that no two events are allowed to happen at the same physical process at the same time—we return to this issue shortly). The task of a logical process is, similarly to the sequential simulator, to remove the first event from the queue, then to update the state of its physical process, and then possibly to send new events to be inserted in the queues of other logical processes. The problem is, of course, that this simple iterative treatment of a logical process’s event queue does not suffice to ensure that, globally, events are processed in nondecreasing order of time. In particular, it is easily conceivable that a logical process receives, for inclusion in its queue, an event scheduled for a time that has already passed in the local simulation.

Classically, the approaches to solving this problem are twofold. Either one makes every effort to ensure that events are indeed globally processed in nondecreasing order of time, or one lets the simulation progress without this constraint and then makes provisions for correcting possible errors when they occur. These two approaches are called, respectively, conservative (discussed in Section 10.3) and optimistic (our subject in Section 10.4), for reasons that come to mind at once.

Before we leave this section, let us discuss in a little more detail the way logical processes interact with one another in both the time-stepped and the event-driven cases. We let $n = \mid \mathcal{N} \mid$, and for $1 \leq i \leq n$ let node $n_i$ be the logical process responsible for simulating the behavior of the physical process $\nu_i$. $G$ is then a graph whose edge set (undirected or directed) has to contain edges at least in one-to-one correspondence to those of $\mathcal{G}$, but $G$ can also be assumed to be a complete undirected graph if the structure of $\mathcal{G}$ is not known. Some of the methods that we discuss require $G$ and $\mathcal{G}$ to be isomorphic, and as a consequence are only applicable to cases in which the structure of $\mathcal{G}$ is known. Other methods do not pose such a requirement, and then it is best to go for generality and assume that the structure of $\mathcal{G}$ is not known, in which case $G$ is taken to be a complete undirected graph.

In a time-stepped simulation by $G$, the messages that nodes exchange are simply the initial or updated states of the corresponding physical processes. In event-driven simulations, a message sent by node $n_i$ to node $n_j$ stands for an event that the physical process $\nu_j$ causes to happen at the physical process $\nu_i$. Such a message is denoted simply by $\text{event}(t)$, where $t$ is the time at which the event is to happen in the physical process $\nu_j$. Depending on the particular application at hand, this message will of course have to contain more information for $n_i$ to be able to simulate the event that it stands for in $\nu_i$. However, in a general framework like ours such details are of little importance, and so we keep the notation conveniently minimal.

Unless we explicitly state otherwise, the distributed simulation that the logical system (that is, $G$) carries out resorts, for termination, to techniques of the sort we discussed in Section 6.2. This is termination in the usual sense, that is, the sense in which all nodes are idle and all edges are empty.

It is very important for the reader to note that the $t$ in the $\text{event}(t)$ message has nothing to do whatsoever with the local time at node $n_i$, in the sense introduced in Section 3.1. This $t$, sometimes also referred to as virtual time, is the time at which the physical process $\nu_j$ at which the event is to happen during the simulation by $n_i$, and is as such part of the model of the natural system that the physical system represents. In spite of such a fundamental difference between the two notions of time, the times at which events are to happen at the various physical processes are also related to each other by restrictive properties, in a way similar to the relation $\prec$ on the set of events in a distributed computation.

One fundamental property says that, if an event is generated at time $t$ by physical process $\nu_j$ to happen at time $t'$ at physical process $\nu_i$, then $t < t'$. This property expresses the unavoidable delay that accompanies the transmission of information between the two physical processes at any finite speed. Another fundamental property, also originating from inherent limitations of physical processes, is that, if $t$ and $t'$ are times at which two events happen at a certain
physical process, then either \( t < t' \) or \( t' < t \). In the remainder of the chapter, we refer to these
two properties as \textit{causality properties}.
The task of an event-driven simulator, be it sequential or otherwise, is to ensure that, from
the perspective of every single physical process, no event is processed unless all events that
precede it in that physical process have themselves already been processed. As we
remarked earlier, a sequential simulator can guarantee this trivially. A distributed simulator,
by contrast, has a whole suite of techniques to choose from with that goal. So an intuitively
appealing interpretation for the differences in the concepts of
the present case the aforementioned causality properties have to be forced upon the logical
system.

In all further sections in this chapter, we let \( \tau_i \) be a variable of node \( n_i \) to contain the time at
physical process \( v_i \). This variable is a nonnegative integer, being initialized to zero. Also, \( n_i \)
employs a variable \( \text{state}_i \), initialized to \( x(0) \), to contain the state of physical process \( v_i \). Unless
otherwise noted, we assume that, when updating \( \text{state}_i \), node \( n_i \) also outputs (to the "user")
the pair \( \tau_i, \text{state}_i \).

\subsection*{10.2 Time-stepped simulation}
Our treatment of time-stepped simulation in this section unfolds along two main
lines, each motivated by a particular class of physical systems. Physical systems in
the first class are said to be \textit{fully concurrent}, while physical systems in the second
class are referred to as being \textit{partially concurrent}. We assume that

\[ G \]

is an
undirected graph, that its structure is known, and that \( G \) and \( \tilde{G} \) are isomorphic
graphs.

In a fully concurrent physical system, \( x(0) \) is provided for every physical process \( v_i \)
and for \( t > 0 \) \( x(t) \) is a function of \( x(t-1) \) and of \( x_j(t-1) \) for every physical process \( v_j \)
such that \( (v_i, v_j) \in \mathcal{E} \). In a partially concurrent physical system, \( x(0) \) is also provided
for every physical process \( v_i \), but the so-called "neighborhood constraints" restrict
the values of \( t \) at which the state of a physical process may be updated.
Specifically, two physical processes \( v_i \) and \( v_j \) can have their states updated for
the same \( t \) if and only if \( (v_i, v_j) \in \mathcal{E} \), that is, \( x_i \) and \( x_j \) do not depend on each other to be
updated. The set of physical processes that can have states updated for the same
\( t \) must then constitute an independent set in \( G \). Because of these neighborhood
constraints, partially concurrent physical systems also come with the requirement
that every node is to be updated infinitely often, which, in our present context,
means the following. If \( t_i \) and \( t_s \) are nonnegative integers such that \( t_s - t_i \geq n - 1 \),
then every physical process \( v_i \) must be such that \( x_i \) is updated for at least one time
\( t \) such that \( t_i \leq t \leq t_s \). Examples of fully concurrent physical systems are cellular automata, various neural
networks, and other systems whose behavior is described by differential equations.
Notorious partially concurrent physical systems include binary Hopfield neural
networks, Markov random fields in general (including Boltzmann machines), and
Bayesian networks. Based on what happens in the cases of these examples, the
time-stepped simulations that we consider in this section terminate when they
converge according to some application-dependent criterion. The value of \( T \) is then
unknown beforehand, and then termination in the sense of idle nodes and empty
edges in the logical system never really occurs. This is the situation we referred to
in Section 6.2 in which it is convenient to employ a leader-based approach to
termination detection. Specifically, every node, upon updating the state of its
physical process, sends a report to the leader with the new state. The leader, in
possession of such reports, is capable of putting together global states (cf. (Section
5.2.2), and on these global states detecting whatever convergence is required for
termination. Upon detection, the leader instructs the other nodes to terminate,
possibly after some additional exchange of messages for gracefulness at the termination. We pursue this termination issue no more in this section, but do nonetheless advise the reader to consider the problem as an exercise (cf. Exercise 1).

Interestingly, already along the book we have seen solutions to the time-stepped simulation of both types of physical system. Specifically, the simulation of fully concurrent physical systems requires a synchronous algorithm that employs pulse \( s = 0 \) for initial states to be exchanged among neighbors and for \( s > 0 \) updates the states of all physical processes for time \( t = s \). Because every state of every physical process has to be sent to all of the corresponding logical process's neighbors, such a simulation is a synchronous algorithm with the property that, at every pulse, exactly one message is sent to every neighbor. Couple this property with the assumption of FIFO edges in \( G \), and we see that Algorithm \( A \_Schedule \_AS \) of Section 5.3.2 can be employed directly for the time-stepped simulation of fully concurrent physical systems (the case of non-FIFO edges is left for the reader's appreciation—cf. Exercise 2).

Before being more specific about the algorithm we have seen that can also be applied to the time-stepped simulation of partially concurrent physical systems, we provide Algorithm \( A \_Simulate \_FC \) (“FC” for Fully Concurrent) for the time-stepped simulation of fully concurrent physical systems based on Algorithm \( A \_Schedule \_AS \). The two algorithms are entirely analogous to each other, so essentially we are simply re-writing the previous algorithm to employ appropriate notation. The variables \( s, \) \( MSG, \) and \( queue_i^j \) for \( n_j \in Neig_i \), employed in Algorithm \( A \_Schedule \_AS \), are now replaced respectively with \( T_i \) (already introduced), a set of variables \( state_i^j \) for \( n_j \in Neig_i \), and \( next \_state_i^j \) both initialized to nil, contain if different from nil the last two states received from \( n_j \) (\( state_i^j \) is the least recent, therefore to be used first).

**Algorithm A\_Simulate\_FC**

**Variables:**
\[
T_i = 0; \\
state_i = x(0); \\
state_i^j = \text{nil for all } n_j \in Neig_i; \\
next \_state_i^j = \text{nil for all } n_j \in Neig_i; \\
initiated = \text{false.}
\]

**Listing 10.1**

**Input:**
\[
msg = \text{nil.}
\]

**Action if \( n_i \in N_0 \):**
initiated, := true;
Send state, to all $n_j \in \text{Neig}_i$. 

**Listing 10.2**

**Input:**
$\text{msg}_i = x$ such that $\text{origin}(\text{msg}_i) = (n_i, n_j)$.

**Action:**
if not initiated then  
begin
  initiated, := true;
  Send state, to all $n_k \in \text{Neig}_i$.
end;
if $\text{state}_i^j \neq \text{nil}$ then
  $\text{next\_state}_i^j := x$
else
  $\text{state}_i^j := x$;
if $\text{state}_i^k \neq \text{nil}$ for all $n_k \in \text{Neig}_i$ then  
begin
  $\tau_i := \tau_i + 1$;
  Update state,;
  Send state, to all $n_k \in \text{Neig}_i$;
  for all $n_k \in \text{Neig}_i$ do
    begin
      $\text{state}_i^k := \text{next\_state}_i^k$
      $\text{next\_state}_i^k := \text{nil}$
    end
  end
end.

Actions (10.1) and (10.2) originate, respectively, from actions (5.19) and (5.20). In this algorithm, the set $N_0$ is the set of nodes that initiate the simulation concurrently. Let us now return to our earlier remark that an algorithm has been seen earlier in this book for the time-stepped simulation of partially concurrent physical systems as well. By definition of a partially concurrent physical system, we see that its time-stepped simulation by the corresponding logical system has to obey the constraints that no two neighbors ever update the states of their physical processes concurrently, and that every node update the state of its physical process infinitely often (in the sense explained earlier). Well, aside from the initial exchange of
states, this is exactly the type of computation that is carried out by Algorithm $A_{Dine\_H}$ of Section 8.3.2.

That algorithm, as we recall, implements scheduling by edge reversal, and is an asynchronous algorithm that functions as follows. Assume that $G$ is initially oriented by an acyclic orientation (cf. Section 8.3.1). Sinks in this orientation must not be neighbors of one another, and may then update the states of their physical processes concurrently. Upon doing such an update, a node sends the new state to all of its neighbors, thereby implicitly reversing the orientation of all edges incident to it and becoming a source. As we remarked in Section 8.3.1, the resulting acyclic orientations are thus guaranteed to be always acyclic, so sinks always exist and the simulation can always progress. In addition, Theorem 8.4 guarantees the form of infinitely often updates we are seeking.

The simulation algorithm based on Algorithm $A_{Dine\_H}$ is given next as Algorithm $A_{Simulate\_PC}$ ("PC" for Partially Concurrent). Unlike the fully concurrent case, now the transformation is a bit more subtle, because initial states need to be spread selectively, that is, node $n_i$ needs the initial state of neighbor $n_j$'s physical process only if $n_i$ is "downstream" from $n_j$ with respect to the initial orientation (i.e., the initial orientation of edge $(n_i, n_j)$ is from $n_j$ to $n_i$). If such is not the case, then the state of that process to be used when $n_i$ computes for the first time will be the one received from $n_j$ upon reversal of the edge $(n_i, n_j)$, so no initial state is really needed. We encourage the reader to write Algorithm $A_{Schedule\_PC}$ in this fashion (cf. Exercise 3), but for simplicity we provide a version in which all nodes send initial states to every neighbor, though at times uselessly.

The additional variables employed by node $n_i$ in Algorithm $A_{Simulate\_PC}$ are, for all $n_j \in Neig_i$, $(initialized\ to\ nil)$ and the Boolean

\[
\begin{align*}
\text{state}_i^j &= \text{nil} \\
\text{upstream}_i^j &= \text{false} \\
\text{holds\_fork}_i^j &= \text{true}
\end{align*}
\]

The variable (which is equivalent to in Algorithm $A_{Dine\_H}$) indicates the current orientation of edge $(n_i, n_j)$, and has to be initialized in accordance with the initial acyclic orientation (to $true$ if the edge is directed from $n_i$ to $n_j$, to $false$ otherwise). The variable

\[
\text{upstream}_i^j = \text{true},
\]

contains the state of $n_j$'s physical process to be used by $n_i$ when it next updates the state of its own physical process.

**Algorithm $A_{Simulate\_PC}$:**

**Variables:**

\[
\begin{align*}
\tau_i &= 0; \\
\text{state}_i &= x(0); \\
\text{state}_j^n &= \text{nil} \ \text{for all } n_j \in Neig_i; \\
\text{upstream}_i^n &= \text{false} \ \text{for all } n_i \in Neig_i.
\end{align*}
\]

**Listing 10.3**

**Input:**

\[
msg_i = \text{nil}.
\]

**Action if $n_i \in N_0$:**
Send state, to all \( n_j \in \text{Neig}_i \).

### Listing 10.4

**Input:**
\[
msg_i = x \text{ such that } \text{origin}(msg_i) = (n_i, n_j).
\]

**Action:**
\[
\text{if } state_i^k = \text{nil} \text{ for all } n_k \in \text{Neig}_i \text{ then}
\]
\[
\text{Send state}_i \text{ to all } n_k \in \text{Neig}_i;
\]
\[
\text{if } state_i^j \neq \text{nil} \text{ then}
\]
\[
\text{upstream}_i^j := \text{true};
\]
\[
\text{state}_i^j := x;
\]
\[
\text{if } (state_i^k \neq \text{nil} \text{ and } \text{upstream}_i^k) \text{ for all } n_k \in \text{Neig}_i \text{ then}
\]
\[
\text{begin}
\]
\[
\text{Update } \tau_i;
\]
\[
\text{Update state}_i;
\]
\[
\text{upstream}_i^k := \text{false} \text{ for all } n_k \in \text{Neig}_i;
\]
\[
\text{Send state}_i \text{ to all } n_k \in \text{Neig}_i;
\]
\[
\text{end.}
\]

In Algorithm \( A\_Simulate\_PC \), actions (10.3) and (10.4) relate closely to actions (8.15) and (8.16) of Algorithm \( A\_Dine\_H \). The differences are accounted for by the need for initial states to be exchanged. Specifically, the set \( N_0 \) of spontaneous initiators no longer corresponds to the initial set of sinks, but rather to the set of nodes that initiate the propagation of initial states spontaneously. In addition, upon receiving a state \( x \) from node \( n_j \), node \( n_i \) must decide whether this is the first message it receives (that is, whether \( state_i^k = \text{nil} \) for all \( n_k \in \text{Neig}_i \)), in which case it must send its physical process’s initial state out. Node \( n_i \) must also be able to distinguish states that it receives from \( n_j \) that should be interpreted as edge reversals (when \( state_i^j \neq \text{nil} \)) from those that are initial values and do not imply edge reversal (when \( state_i^j = \text{nil} \)). Finally, testing whether \( n_i \) has become a sink requires not only that \( state_i^k \) be checked for all \( n_k \in \text{Neig}_i \), but...
as well (because if \( n_i \) is an initial sink, it may only update \( state_i \) after receiving initial states from all of its neighbors).

As a final observation, we note that the way to update \( T_i \) in (10.4) was left purposefully vague. It should be clear that simply adding one as in (10.2) does not suffice, because it is necessary to account for all the time units in which \( state_i \) was not updated because \( n_i \) was not a sink. We leave it to the reader to remove this vagueness (cf. Exercise 4).

### 10.3 Conservative event-driven simulation

In this section, we elaborate on conservative approaches to distributed event-driven simulation. As we remarked in Section 10.1, conservative methods seek to guarantee that events are processed in increasing order of time at all nodes, so that, globally, events are processed in nondecreasing order of time.

Our treatment of conservative methods is presented in the two sections that follow. The method that we present in Section 10.3.1 requires isomorphism between \( G \) and \( \dot{G} \), and in addition that the edges of \( \dot{G} \) be FIFO. Such a method is then only applicable to cases in which the structure of \( \dot{G} \) is known. In Section 10.3.2, by contrast, a conservative method is discussed that does not require such an isomorphism, and consequently is applicable to cases in which the structure of \( \dot{G} \) is not known. In these cases, \( G \) is taken to be a complete undirected graph, but the method still requires FIFO edges in \( G \).

In addition to the causality properties that we discussed in Section 10.1, in the case of conservative methods another property is needed on the physical system. If a sequence of \( k \geq 1 \) events is generated at times \( t_1 < t_2 < \ldots < t_k \) by a physical process \( \nu_i \) to happen respectively at times \( t'_1, t'_2, \ldots, t'_k \) at physical process \( \nu_j \), then we require that

\[
0 \leq t'_1 < t'_2 < \ldots < t'_k
\]

as well. This is a monotonicity property. Note that, although conservative methods seek to faithfully mimic the functioning of a sequential simulator, in the sequential case monotonicity is not an issue (i.e., it does not need to happen) so long as the causality properties hold. Requiring monotonicity in the conservative distributed case may be thought of as requiring that \( \dot{G} \)'s edges, like those of \( G \), be FIFO.

#### 10.3.1 A first algorithm

For greater generality, in this section we take \( \dot{G} \) (and consequently \( G \)) to be a directed graph. In addition, if \( \dot{G} \) is not strongly connected, then we assume that all of its sources are in \( \mathcal{N} \).

The monotonicity property of the physical system and the assumption that \( G \)'s edges are FIFO guarantee that, if a node processes \( event(t) \) messages that it receives in increasing order of \( t \), then the sequence of \( event(t) \) messages that it sends to another node \( n_i \) is received in increasing order of \( t \) as well. In order to guarantee that \( n_i \) too processes the \( event \) (\( t \) messages that it receives in increasing order of \( t \), by the causality properties all \( n_i \) has to do is to merge the incoming streams of \( event(t) \) messages so that the resulting single stream is sorted in increasing order of \( t \). It then suffices for \( n_i \) to process \( event \) messages as they are queued in this resulting stream.

The approach that suggests itself for the participation of node \( n_i \) in the simulation is then the following. For each \( n_j \in I_{\text{Neig}_i} \), node \( n_i \) does not really maintain a queue of incoming messages, but rather a variable \( next_time_j^i \), initialized to zero, to contain the \( t \)
parameter in the next event(t) message to be processed in the stream from \( n_i \) (if no such

next message has been received, then \( \text{next\_time}_{i}^{k} \) is either the \( t \) in the last event

(t) message received from \( n_i \) or the initial zero). All \( n_i \) has to do is then to select for

processing the event message from \( n_k \in I_{Neig} \), such that

\[
\text{next\_time}_{i}^{k} < \text{next\_time}_{i}^{j}
\]

for all \( n_j \in I_{Neig} \), such that \( n_i \neq n_k \). After processing this event message, \( n_i \) waits to receive

another event message from \( n_i \), so that a new minimum can be found.

The problem with this initial approach is of course that such a message may never be

received, and then the simulation may deadlock. In fact, this same problem exists from the

very beginning of the simulation, because no node \( n_i \) may process any event message before

receiving one message from each neighbor in \( I_{Neig} \). As in general situations in which

deadlocks may happen, here too we have a choice as to either prevent its occurrence or

detect it after it occurs (cf. Section 6.3). We chose the prevention strategy, but approaches

based on deadlock detection have also been proposed.

The deadlock-prevention fix that we add to the simulation strategy we just described is based

on additional null(t) messages that the nodes exchange. Such messages are sent by \( n_i \) to
every neighbor in \( O_{Neig} \), to which it does not send an event upon computing the

aforementioned minimum and simulating the behavior of \( v_i \) up to that minimum. If the

message that corresponds to this minimum time is itself a null message, then a null is sent to

all of \( n_i \)'s neighbors in \( O_{Neig} \). Although null messages may account for an excessive

increase in the algorithm’s message complexity, they can also be used to advance the

simulation more rapidly if so the physical system's peculiarities permit, and to enable nodes
to terminate their computations locally.

Before proceeding to the presentation of the algorithm, let us be more specific about these
alternative uses for a null message. When an event(t) or a null(t) is processed by node \( n_i \), a
null(t + 1) message is sent to every node \( n_k \in O_{Neig} \), to which an event message is not

sent. This null(t + 1) message has the purpose of informing \( n_i \) that it will never be sent by \( n_i \)

any event(t') message such that \( t' \leq t + 1 \).

If the physical process being simulated is such that \( n_i \) can predict that no event(t') message
will ever be sent to \( n_i \) such that \( t' > t + 1 \), then a null(t + 1) message is sent instead.
This message may be quite beneficial to the processing at \( n_i \), in the sense that it may allow \( n_i \)
to make more progress before it has to wait for additional messages from \( n_i \). The difference

\( t' - t - 1 \) is known as \( n_i \)'s lookahead at time \( t \) with respect to node \( n_i \). In the algorithm to be

presented, and for all \( n_j \in O_{Neig} \), \( n_i \) maintains a variable \( \text{lookahead}_{i}^{j} \), which we
assume is properly initialized and maintained based on the physical system’s characteristics.
Similarly, when the event(t) message that \( n_i \) must send is such that \( t > T \), then a null(T) is
sent instead, thereby signaling the destination node that \( n_i \) is never going to send it another

message. This is also what nodes that correspond to sources in \( G \) must do after

participating in the algorithm for time zero.

The classical example of a physical system exhibiting the possibility of lookahead
determination is that of a network of queues (although it is easy to argue that in general
simulating such a system by a distributed algorithm is not a good idea—reliable results for
such a physical system require multiple simulations on random initial conditions, and then it
is best to employ multiple sequential simulations). In such a system, each physical process is
a queue with a server that provides service to each customer in the queue according to a
certain distribution of service times. If for a queue it holds that no service time is ever less
than, say, \( z \) time units, then the lookahead of the corresponding logical process is at all times
equal to \( z \) with respect to all of its neighbors to which it may send event messages.

In order to contain the number of null messages when they are also used for lookahead and
termination purposes, node \( n_i \) keeps track, for all \( n_j \in O_{Neig} \), of the value of \( t \) in the last
null(t) message sent to \( n \). Variable, initially set to zero, is employed for this purpose.

The algorithm that realizes this conservative simulation strategy is Algorithm `A_Simulate_C` ("C" for Conservative), presented next. In addition to the already introduced variables, it also employs the following. For \( n_j \in I_{Neig} \), the Boolean initialized to true, indicates whether \( n \) must receive a message from \( n_j \) before continuing. Also, the Boolean indicates, if \( = \) false, whether the message corresponding to \( i \) was a null message. Finally, the auxiliary set \( X_i \) and variable \( t \) are also employed.

**Algorithm A_Simulate_C:**

Variables:

\[
\begin{align*}
\tau_i &= 0; \\
state_i &= x(0); \\
next_time_i &= 0 \text{ for all } n_i \in I_{Neig}; \\
next_time_i &= 0 \text{ for all } n_i \in O_{Neig}; \\
needs_event_i &= \text{true} \text{ for all } n_i \in I_{Neig}; \\
null_message_i &= \text{false} \text{ for all } n_i \in I_{Neig}; \\
X_i &= \text{for all } n_i \in I_{Neig}; \\
t_i &= \text{for all } n_i \in I_{Neig};
\end{align*}
\]

**Listing 10.5**

Input:

\[ \text{msg}_i = \text{nil.} \]

Action if \( n_i \in N_0 \):

Let \( X_i \subseteq O_{Neig} \) be the set of nodes to which event's are to be sent;

for \( n_j \in X_i \) do begin

Let \( \tau_i \) be the time of the event to be sent to \( n_j \);

if \( \tau_i \leq T \) then

Send event(\( \tau_i \)) to \( n_j \);
if $t > T$ or $I_{\text{Neig}} = \emptyset$ then
Send $\text{null}(T)$ to $n_j$
end;

for $n_i \notin X$ do
begin

\[
\text{last}_i := \min(\text{last}_i, \text{lookahead}_i, T);
\]
Send \text{null(last)} to $n_j$
end.

\underline{Listing 10.6}

Input:
$\text{msg}_i = \text{event}(t)$ such that $\text{origin}(\text{msg}_i) = (n_j \rightarrow n_i)$.

Action when
\[
\text{needs}_i^j := \text{false};
\text{null}_i^j := \text{false};
\text{next}_i^j := t;
\text{needs}_i^k := \text{false};
\]

if \textbf{not} for all $n_k \in I_{\text{Neig}}$, then
begin
Let $n_k \in I_{\text{Neig}}$ be such that all $n_i \in I_{\text{Neig}}$;
\[
\text{next}_i^k \leq \text{next}_i^j
\]
if then $X_i := \emptyset$
else
begin
Update \text{state};
Let $X_i \subseteq O_{\text{Neig}}$ be the set of nodes to which event's are to be sent
end;
end;
for $n_i \in X$ do
begin
Let $t > \text{next}_i^j$ be the time of the event to be sent to $n_i$;
if $t \leq T$ then
Send \text{event}(t) to $n_i$
else
    Send $null(T)$ to $n_i$
end;
for $n_i \notin X$ do
begin

\[
\text{last_time}_i^l < \min\{ \text{\tau}_i^l + 1 + \text{lookahead}_i^l, T\} \text{ then}
\]
begin

\[
\text{last_time}_i^l := \min\{ \text{\tau}_i^l + 1 + \text{lookahead}_i^l, T\};
\]
Send $null(\text{last_time}_i^l)$ to $n_i$
end
end;
\[
\text{needs_event}_i^k := \text{true}
\]
end.

Listing 10.7

Input:
$\text{msg}_i = \text{null}(t)$ such that $\text{origin}(\text{msg}) = (n_j \rightarrow n_i)$.

Action when $\text{\text{needs_event}}_i^j := \text{false}$;
$\text{null_message}_i^j := \text{true}$;
$\text{next_time}_i^j := t$;
$\text{needs_event}_i^k$ for all $n_k \in I_{\text{Neig}}$, then

Let $n_k \in I_{\text{Neig}}$ be such that all $n_k \in I_{\text{Neig}}$;
\[
\text{\tau}_i^l := \text{next_time}_i^k
\]
if $\text{null_message}_i^k$ then $X_i := ()$
Update state;
Let \(X_i \subseteq O_{\text{Neig}}\) be the set of nodes to which event's are to be sent
end;
for \(n_i \in X\) do begin
Let \(T_i\) be the time of the event to be sent to \(n_i\);
if \(T_i \leq T\) then
Send event\((T)\) to \(n_i\)
else
Send null\((T)\) to \(n_i\)
end;
for \(n_i \notin X\) do begin
if \(\text{last_time}^l_{i} < \min\{T_i + 1 + \text{lookahead}^l_{i}, T\}\) then
begin
\(\text{last_time}^l_{i} := \min\{T_i + 1 + \text{lookahead}^l_{i}, T\};\)
Send null\((\text{last_time}^l_{i})\) to \(n_i\)
end;
\(\text{needs_event}^k_{i} := \text{true}\)
end.

In Algorithm A_Simulate_C, the set \(N_0\) is the set of all nodes that initiate the simulation spontaneously. These nodes include those whose physical processes are in \(N_{0}\), but are not restricted to them (if \(n_i \notin N_0\) but \(v_i \notin N_{0}\), then in (10.5) only null messages are sent). Action (10.5) in the algorithm is executed by the nodes in \(N_0\), while actions (10.6) and (10.7) are executed upon receipt by \(n_i\) of an event message or a null message, respectively. Actions (10.6) and (10.7) are identical to each other, except for the setting of variable \(n_k\), \(n_i\) being the message's origin. It is important to note that, in the algorithm, the causality and monotonicity properties are only implicitly ensured, depending essentially of the portions of the simulation that relate to the nature of the physical system under consideration. Another important observation is that, although in (10.6) and (10.7) the determination of \(n_k\) would be unique in the absence of null messages (by the causality properties), when such messages are employed it may happen that the minimum \(n_k\) occurs for more than one node in \(I_{\text{Neig}}\). However, at most one such node participates in the minimum with an event message instead of a null message, and it is to
such a neighbor that $n_k$ must be preferably set if it exists (the reader should try to be convinced that, if this is the case, then the null messages for which the minimum time also holds will not generate the sending of any messages when they are processed).

We now turn to establishing the algorithm's correctness.

**Theorem 10.1.**
*Algorithm A_Simulate_C correctly simulates the physical system for all $t \in \{0, \ldots, T\}$.*

**Proof:** Every node corresponding to a physical process in $N_0$ is in the set $N_0$ of spontaneous initiators. In addition, by (10.5) through (10.7) a node never sends any event($t$) or null($t$) message for which $t > T$. Because the causality and monotonicity properties are guaranteed to hold and all of $G$'s edges are FIFO, what we need to show is that the simulation always progresses so long as there exists at least one node $n_i$ for which $T_i < T$.

Suppose, to the contrary, that a deadlock happens. It must then be, following our discussion in Section 6.3, that a wait cycle exists in $G$. In this cycle, every node $n_i$ is precluded from picking an event or null message to process because there exists $n_j \in I_{\text{Neig}}$ such that $\text{needs\_event}_i = \text{true}$. Because $N_0$ is nonempty, at least one such wait cycle has to exist including at least one node $n$, such that either $n_i \in N_0$ or $T_i > 0$. The message that this $n_i$ needs from the corresponding $n_j$ in order to continue must carry a parameter $t$ such that $t > T_i$ when it is sent (by the causality properties), which means that there exists $n_k$ in the cycle waiting for a message from $n_i$ with parameter $t$ such that $t > T_i$. But either by (10.5) (if $n_i \in N_0$) or by (10.6) and (10.7) (if $T_i > 0$), such a message must have been sent, respectively spontaneously or when $n_i$ last updated. It is then impossible for any such wait cycle to exist, thence any deadlock as well.

**10.3.2 Conditional events**

The need for the structure of $G$ to be known and the potentially excessive traffic of null messages in Algorithm A_Simulate_C have led to the search for other conservative methods. As we mentioned earlier, some of the other methods that have been proposed are based on the use of deadlock detection, instead of prevention, although in them the need to know the structure of $G$ still persists. In this section, we do not assume that the structure of $G$ is known, and then take $G$ to be a complete undirected graph. $G$'s edges are still assumed to be FIFO edges.

The approach that we discuss in this section is based on the following observation. In a sequential simulation, every event in the single queue of events is a conditional event, in the sense that it must only be scheduled to happen when it reaches the head of the queue, at which time it becomes a definite event. In Algorithm A_Simulate_C, definite events were determined by restraining the input of messages to a node. In other words, only upon having received exactly one message from each neighbor did a node choose from those messages one to be processed. If the chosen message was an event message, then event messages corresponding to definite events were output.

The method to turn conditional events into definite events is in this section different. The messages exchanged among nodes are still event and null messages, but the latter no longer have a deadlock-prevention connotation, but rather are only used to convey lookahead and termination information. Node $n_i$ no longer restrains the receipt of messages, but rather assumes that every incoming event message corresponds to a definite event and because every edge is FIFO they may therefore be acted upon immediately. In order for this
assumption to be valid globally, \( n_i \) makes sure that only \textit{event} messages that correspond to definite events are output. To this end, \( n_i \) computes on every \textit{event} or \textit{null} message it receives, and as a result produces as many \textit{event} and \textit{null} messages as it can. The \textit{null} messages it produces, having solely a lookahead- or termination related meaning, are immediately output when they are generated. Messages of the \textit{event} type, however, are stored in a set \( \text{events}_i \) until the events to which they correspond can become definite, at which time the messages are sent out. At all times, a variable \( \text{next}_i \) indicates the least time \( t \) associated with the \textit{event}(\( t \))'s in \( \text{events}_i \).

Determining which of the \textit{event} messages are to be sent out as definite events (and when this is to happen) is the crux of the approach, and is achieved as follows. For all \( n_k \in N \), node \( n_i \) maintains a collection of variables \( \text{next}_i \) to contain local views of \( \text{next}_k \). For each such \( n_k \), node \( n_i \) also maintains the variables \( \text{sent}_i^k \) and \( \text{received}_i^k \) to indicate the number of messages (of either type) it ever sent to \( n_k \) or received from \( n_k \), respectively. Similarly, for all ordered pairs of nodes \( (n_k, n_l) \), a collection of variables \( \text{sent}_i^{kl} \) and \( \text{received}_i^{kl} \) is also maintained to indicate \( n_i \)'s views of the number of messages sent by \( n_k \) to \( n_l \) and received by \( n_k \) from \( n_l \), respectively.

In order to describe the computation that takes place on these variables, we need to resort to the terminology of Section 5.2.1, where we described an algorithm for global state recording on a substrate computation. In the present case, we regard as a substrate computation the computation that the nodes perform that is related to the simulation proper. That is, \( n_i \)'s participation in the substrate is to compute on \textit{event} and \textit{null} messages it receives, thereby updating all the variables involved, including the variables \text{events}_i, \text{next}_i, the \text{sent}_i^k's, and the \text{received}_i^k's. On top of this substrate computation, a global state recording is performed periodically. The goal of these recordings is to seek global states in which all edges are empty, and then Algorithm \text{A_Record_Global_State} of Section 5.2.1 offers no help. Instead, we recall our observation in Section 5.2.2 on the possibility of recording such global states by a leader, except that no leader will in this case be employed, but rather the recording will be done by all nodes acting as "leaders."

The following is then how the substrate and the global state recording interact. The local state to be recorded at node \( n \) comprises the variables \( \text{next}_i \), the \( \text{sent}_i^k \)'s, and the \( \text{received}_i^k \)'s. These variables must be recorded (respectively in available \( \text{sent}_i^k \)'s, and \( \text{received}_i^k \)'s) from time to time, specifically a finite time after one of them changes. In addition, a finite time after the recording they must be broadcast to all other nodes. Node \( n_i \), now in its role as the aforementioned "leader," collects such broadcasts in available \( \text{received}_i^k \)'s, and looks for system states in which \( \text{sent}_i^{kl} = \text{received}_i^{kl} \) for all ordered pairs \( (n_k, n_l) \). In these system states, all edges are empty, so they must constitute global states, as we observed in Section 5.2.2. If such a global state is detected at which \( \text{next}_i^k \leq \text{next}_i^{kl} \) for all \( n_k \in N \), then \( n_i \) is sure never to add another \textit{event}(\( t \)) such that \( t \leq \text{next}_i^k \) to \text{events}_i. Consequently, every \textit{event}(\( t \)) \( \in \)
events, such that \( t = \text{next}_i^t \) is seen to correspond to a definite event, and may be sent to its destination, while states, and next, are updated accordingly.

We do not provide any further details on this algorithm, but rather leave providing such details for the reader to undertake as an exercise (cf. Exercise 5).

As a final remark, we note that, in addition to the overall scheme for turning conditional events into definite events we just described, node \( n \) may at any time detect, based on specifics of the application at hand, that certain conditional events are in fact definite and may be sent out without waiting for any global information. This is valid for sequential simulations as well (an event that is not at the head of the queue may, depending on the application, be processed), although it makes little sense in that case.

### 10.4 Optimistic event-driven simulation

Optimistic methods of distributed simulation are based on the premise that it may be more efficient to let causality errors occur and then fix them than to rely on lookaheads and other application-specific properties in the search for efficiency. Physical systems for which this premise has proven valid include systems of colliding particles and evolving populations. In this section, then, there is no place for such things as lookaheads and null messages.

Similarly, the structure of \( G \) is not assumed to be known, so that \( G \), whose edges no longer have to be FIFO, is taken to be a complete undirected graph.

The mechanism that we describe in this section for optimistic distributed simulation is known as the time warp mechanism, perhaps in allusion to the possibility that, at node \( n \), \( \tau_i^t \) may move back and forth as the need arises for errors to be corrected. The essence of this mechanism is the following. Whenever \( n \) receives an event message such that \( t > \tau_i^t \), it sets \( \tau_i^t \) to \( t \), computes on the message it received, and possibly sends out event message(s) for some \( t > \tau_i^t \). Because no precautions are taken to ensure that such events are definite (in the terminology of Section 10.3.2), it may well happen that a event message reaches \( n \) with \( t \leq \tau_i^t \), thereby indicating that whatever state updates were done or event messages were sent in the interval \( \{ t, \ldots, \tau_i^t \} \) were erroneous and must therefore be corrected. This arriving event message is often referred to as a "straggler."

The approach of the time warp mechanism to correcting such errors when they are detected is to return the simulation globally to a correct global state, and then to proceed from there. In order to be able to perform such "rollbacks," every node must store some of its past history, so that earlier states can be restored when necessary. At node \( n \), this history has two queue components, called state_queue and output_queue. An element of state_queue is the pair \((t,x)\), indicating the state \( x \) of the physical process \( v \) at time \( t \). This queue is initialized to \( \text{null} \), and receives a new pair whenever state, is updated. An element of output_queue is the triple \((t,t',n)\), indicating that \( n \) sent \( n \), an event message when \( \tau_i^t \) was equal to \( t \). This queue is initialized to \( \text{null} \), and receives a new triple whenever \( n \) sends an event message. Both queues are kept in increasing order of \( t \) (nondecreasing for output_queue, for there may be multiple event's sent for fixed \( \tau_i^t \).

When a straggler arrives with a \( t \) parameter, \( \tau_i^t \) is set to \( t \) and state, is set to \( x \) in the \((t',x)\) pair in state_queue. Here \( t' \) is the greatest integer less than \( t \) for which a pair exists in state_queue. This queue is then shortened to contain pairs with time components no greater than \( t \). Before resuming normal processing, however, \( n \) must have the effect of every event it sent when \( \tau_i^t = t \) or later. This is achieved by sending an anti-event message to the \( n \) in every triple \((t',t',n)\) in output_queue, such that \( t' \geq t \), and then shortening the queue by the removal of those triples. It only remains for us to discuss how to handle the reception of such anti-event's.
Because $G$’s edges are not assumed to be FIFO, an \textit{anti-event}$(t)$ arriving at node $n_i$ may be following the \textit{event}$(t)$ to which it corresponds or it may be ahead of the \textit{event}$(t)$. In the former case, the \textit{anti-event} is also a straggler upon arrival, and should be treated as we discussed previously. In the latter case, $n_i$ needs a mechanism to remember the arrival of the \textit{anti-event} $(t)$, so that, when the \textit{event}$(t)$ arrives, it is not acted upon. In order to implement this mechanism, node $n_i$ maintains yet another queue, called $input\_queue_i$ and initialized to nil, where the pair $(t,n_k)$ corresponding to an \textit{anti-event}$(t)$ from $n_k$ that does not arrive as a straggler is stored in increasing order of $t$. An arriving \textit{event}$(t)$ from node $n_k$ that finds the pair $(t,n_k)$ in $input\_queue_i$ is rendered ineffective, while the queue is shortened by the removal of that pair. (Note that, by the causality properties, only the $t$’s would have to be stored in $input\_queue_i$; however, \textit{anti-message’s} occur in erroneous situations, thence the additional precaution of storing the messages’ origins as well.)

This strategy is realized by Algorithm A\_Simulate\_TW (“TW” for Time Warp), given next. In addition to the variables already described, the algorithm also employs the auxiliary variable $t_i$. Comparing with our initial approach in Section 10.1, no queue of \textit{event} messages to be processed is really needed. Instead, node $n_i$ in Algorithm A\_Simulate\_TW acts upon such messages as they are received.

\textbf{Algorithm A\_Simulate\_TW:}

\begin{verbatim}
Variables:
$\tau_i = 0$;
state, = $x(0)$;
state_queue, = nil;
output_queue, = nil;
input_queue, = nil;
t_i.

Listing 10.8

Input:
$\text{msg}_i = \text{nil}$.  
Action if $n_i \in N_0$:  
  Append $(\tau_i, \text{state},)$ to state_queue,;  
  for $n_j \in \text{Neig}$, do  
    if there exists event to be sent to $n_j$ then  
      begin  
        Let $t > \tau_i$ be the time of the event to be sent to $n_j$;  
        if $t \leq T$ then  
          Send $\text{event}(t)$ to $n_j$  
          end.  
  end.  

Listing 10.9

Input:
$\text{msg}_i = \text{event}(t)$ such that $\text{origin}(\text{msg}_i) = (n_i, n_j)$.  
Action:  
  if there exists $(t, n_j)$ in input_queue, then  
    Remove $(t, n_j)$ from input_queue,;  

end.
\end{verbatim}
else
begin
if \( t \leq T_i \) then
begin
Let \( t' \) be the greatest integer such that \( t' < t \) and there exists \((t', x)\) in \text{state\_queue};
state := x;
Remove all \((t', x')\) such that \( t' \geq t \) from \text{state\_queue};
for all \((t', t', n_i)\) in \text{output\_queue} such that \( t' \geq t \) do
begin
Remove \((t', t', n_i)\) from \text{output\_queue};
Send anti-event(t') to \( n_i \)
end;
end;
\( T_i := t; \)
Update \text{state};
Append \((T_i, \text{state})\) to \text{state\_queue};
for \( n_i \in \text{Neig} \) do
if there exists event to be sent to \( n_i \) then
begin
Let \( t > T_i \) be the time of the event to be sent to \( n_i \);
if \( t \leq T \) then
begin
Append \((T_i, t, n_i)\) to \text{output\_queue};
Send event(t) to \( n_i \)
end;
end.
end.

\underline{Listing 10.10}

\textbf{Input:}
\( \text{msg}_i = \text{anti-event}(t) \) such that \( \text{origin}(\text{msg}_i) = (n_i, n_j) \).

\textbf{Action:}
if \( t \leq T_i \) then
begin
Let \( t' \) be the greatest integer such that \( t' < t \) and there exists \((t', x)\) in \text{state\_queue};
state := x;
Remove all \((t', x')\) such that \( t' \geq t \) from \text{state\_queue};
for all \((t', t', n_i)\) in \text{output\_queue} such that \( t' \geq t \) do
begin
Remove \((t', t', n_i)\) from \text{output\_queue};
Send anti-event(t') to \( n_i \)
end;
The set \( N_0 \) in Algorithm \( A_{Simulate_TW} \) comprises the nodes whose physical processes are in \( N_0 \). Action (10.8) corresponds to the processing on events at physical processes in \( N_0 \), while actions (10.9) and (10.10) correspond, respectively, to the receipt at \( n_i \) of an event message and an anti-event message. Both actions include \( n_i \)'s participation in a rollback if the message is a straggler. In the case of an event message, this participation in the rollback is followed by the processing of the event. If the message is not a straggler, then the corresponding event is processed in (10.9) or a new element is added to \( input_queue_i \). The reader will have noticed that the sending of event's in (10.8) and in (10.9) differ from each other in that (10.8) does not include any additions to \( output_queue_i \). As a result, \( output_queue_i \) does not contain any pair \((0, t', n_k)\). What this amounts to is that provisions are not made for a possible rollback of the simulation in which \( n_i \) must return to \( =0 \). The reason why such provisions are indeed unnecessary, and in fact why numerous other properties of Algorithm \( A_{Simulate_TW} \) and variations thereof hold, relies on the following definition. At any global state, consider the minimum of the following quantities: for all \( n_i \in N \) and for every message in transit sent by node \( n_i \) the value of \( at \) the moment the message was sent. This minimum is called the \textit{global virtual time} at that global state, known mainly by the acronym \( GVT \) (for Global Virtual Time).

**Theorem 10.2.**

\textit{In Algorithm} \( A_{Simulate_TW} \), \( \tau_{i} \) is never set to a value \( t \leq GVT \).

\textbf{Proof:} The value of \( \tau_{i} \) is only changed to \( t \) upon receipt of an event \((t)\) in (10.9) or an anti-event \((t)\) in (10.10). By the physical system's causality properties, any such message, when sent by \( n_k \in N \), must have been sent when \( \tau_{k} < t \). The theorem then follows from the observation that \( GVT \leq \tau_{k} \) (and consequently \( GVT < t \)) at any global state in which the said message is in transit.

At the initial global state of the simulation, \( GVT = 0 \). The reason why \( output_queue_i \) does not contain any elements with a zero time component for any node \( n \) is then immediate from Theorem 10.2. This theorem, in addition, implies the following.

**Corollary 10.3.**

\textit{Algorithm} \( A_{Simulate_TW} \) \textit{correctly simulates the physical system for all} \( t \in \{0,...,T\} \).

\textbf{Proof:} This is an immediate consequence of the fact that every node whose physical process is in \( N_0 \) is in \( N_0 \), the fact that in (10.8) and in (10.9) event(t)'s are never sent with \( t \) \( T \), and the physical system's causality properties, if only we consider that, by Theorem 10.2, progress in the simulation is always guaranteed.

In addition to being instrumental in establishing the correctness of Algorithm \( A_{Simulate_TW} \), the \( GVT \) concept is also useful in other situations, including memory management at the various nodes. Specifically, the only pairs that need to be maintained in \( state_queue \), at any node \( n \), are those with time component \( t' \geq t \), where \( t \) is the greatest
integer such that \( t \leq GVT \) for which a pair exists in \( state\_queue \). Similarly, \( output\_queue \) need not contain any \((t, t', n)\) for \( t \leq GVT \). These are immediate consequences of Theorem 10.2.

When employed for such memory management purposes, the value of \( GVT \) needs from time to time to be accessible locally to the nodes. Regardless of which technique is employed for this to happen (either a global state recording algorithm in the fashion of Section 5.2.1 or some of the other techniques present in the literature requiring fewer messages), \( event \) and \( anti-event \) messages can no longer be sent as we have introduced them, but instead must include another time parameter to store the value of the variables at the time they are sent. So these messages must then be sent by \( n_i \) as \( event(\bar{T_i}, t) \) and \( anti-event(\bar{T_i}, t) \), respectively.

Let us make one final observation before leaving this section. As we remarked in Section 10.1, the updating by node \( n_i \) of \( state_i \) is implicitly taken as also implying that the pair \( (\bar{T_i}, state_i) \) is output. It should be clear to the reader that guaranteeing this in Algorithm \( A\_Simulate\_TW \) requires a little more elaboration. Specifically, such a pair can only be output if \( \bar{T_i} \leq GVT \), thereby providing another justification for the need to acquire estimates of \( GVT \) locally from time to time.

10.5 Hybrid timing and defeasible time-stepping

There are physical systems in which the states of the physical processes change in a way that does not entirely fall into any of the two categories we introduced in Section 10.1. For such systems, the methods we have seen so far in the book are inadequate, and other alternatives have to be devised. In this section, we briefly describe an example of such physical systems and outline a simulation strategy that can be regarded as a hybrid between some of the approaches we have studied. The physical system we describe arises from problems in nuclear physics, and it appears that some phenomena associated with the dynamics of stellar cores can be modeled likewise.

The physical system consists of \( p \) interacting particles in three-dimensional space. At time \( t \geq 0 \), the particles' positions are \( z_1(t), \ldots, z_p(t) \) and their momenta are \( p_1(t), \ldots, p_p(t) \). All particles have mass \( m \), and their behavior can be modeled by integral-differential equations whose solution cannot be obtained analytically or even numerically within reasonable bounds on the required computational resources. The approach to solving them is then to employ a heuristic that assumes simpler modeling equations and uses randomness to guarantee accuracy. We describe such a heuristic next.

For \( 1 \leq k \leq p \), the behavior of the \( k \)th particle is assumed to follow the equations

\[
\frac{dz_k(t)}{dt} = \frac{p_k(t)}{m},
\]

\[
\frac{dp_k(t)}{dt} = -\nabla U\left(\rho(z_k(t))\right),
\]

where \( U \) is a potential and \( \rho(z_k(t)) \) is the average particle density in the vicinity of point \( z_k(t) \) at time \( t \). The average here is the average over a large number, call it \( N \), of random initial conditions, so what is required is the solution of \( Np \) pairs of differential equations like the ones we showed. These equations are very tightly coupled with one another, so what we have is not the typical situation in which \( N \) independent solutions are required, in which case distributed methods can hardly be recommended within the context of obtaining each of the individual solutions. Instead, our problem is to solve for the positions and momenta of \( Np \) interacting particles, based on equations that require knowledge, at all times, of the average particle density, over \( N \), near the particles' positions.
Because analytical methods to solve this system of equations are not known either, the sequential approach is to employ simulation. For a conveniently chosen $\Delta t$, the simulation starts at initial positions and momenta for all $N_p$ particles and computes these quantities for the discrete times $\Delta t, 2\Delta t, \ldots, T$. It is, in this sense, a time-stepped simulation. If $t$ is any of these discrete times, then the solutions at $t$ are computed from the solutions at $t - \Delta t$ and from the average densities corresponding to the interval $[t - \Delta t, t)$. The problem, naturally, is the computation of such densities, because they depend on how the particles interact with one another during that interval. What is done in the sequential method is to perform an event-driven simulation for each of the intervals, with provisions for events not to be generated for occurrence at times $t'$ such that $t' \geq t$.

This hybrid sequential method has an obvious distributed counterpart, which is the following. The time-stepped portion can be achieved by a synchronous algorithm. By means of any of the synchronizers seen in Section 5.3, this synchronous algorithm can be turned into an asynchronous one. The processing for each pulse is an event-driven simulation that must terminate before nodes are allowed to progress to further pulses. This event-driven simulation can employ any of the approaches we discussed in Sections 10.3 and 10.4, for example.

An alternative that is not so tightly synchronized is to employ essentially the same guidelines we just described, with the slight modification that an optimistic method be used within each pulse, and that pulses, like events in the optimistic simulation, be defeasible, in the sense of being prone to annulment by way of rollbacks. Let us be a little more specific on a method, called defeasible time-stepping, that proceeds along these lines.

We present the method's essentials for the case of the physical system introduced at the beginning of this section. A physical process is a region in the portion of three-dimensional space to which the particles are confined, so the overall structure of $\mathcal{G}$ is quite well known. Within each time interval, the events that characterize the interaction among the physical processes are the arrival of particles from neighboring regions in space, so $\mathcal{G}$ is an undirected graph.

Defeasible time-stepping operates on a graph $G$ that is isomorphic to $\mathcal{G}$ for this particular problem.

In order to carry out the event-driven simulation corresponding to each time interval, a logical process requires the particle densities at points within its own physical process and at points in adjacent physical processes that are in the vicinity of the boundary between the two. What this implies is that logical processes must, at the end of each time interval, send information on the appropriate densities to all of their neighbors in $G$. Clearly, then, the synchronization method that is best suited is the particularization of Synchronizer Alpha that we saw in Section 5.3.2 (i.e., Algorithm A_Schedule_AS), provided edges are FIFO, as we do assume for the sake of simplicity. So the time-stepped portion of the simulation is in principle performed in much the same lines as those of Section 10.2.

A logical process starts participating in the event-driven simulation that corresponds to a time interval as soon as it receives the necessary densities from all of its neighbors corresponding to the previous time interval. This event-driven simulation is performed optimistically, say by means of techniques similar to the one introduced in Section 10.4. When a logical process judges that it may have finished participating in the simulation for the current time interval, it sends densities out to its neighbors. If it then receives a straggler to be processed in that same time interval (or in an earlier one), it must correct the effects of the premature signal for its neighbors to proceed to the next time interval. Overall, events and messages carrying densities are processed optimistically, and may then have to be revised upon the occurrence of errors. The techniques employed to this end are pretty similar to those of Section 10.4. We provide no further details here, but encourage the reader to seek additional information in the pertinent literature.
10.6 A general framework

In this section, we provide a brief description of a framework that may help visualize how all the different approaches to the simulation of physical systems relate to one another. The overall goal of such a unified understanding has some obvious aesthetic connotations, but more importantly is to provide an intuitive basis for the development of new methods for physical systems with new characteristics.

In this unifying framework, the physical system is viewed as a two-dimensional grid, with one dimension (say the "horizontal") used to represent the physical processes and the other (the "vertical") used to represent time. Each point in this grid corresponds to a physical process and a time instant. The goal of a simulation method is to fill out the grid by assigning to each point the state of the corresponding physical process at the corresponding time.

In broad terms, a logical process may correspond to any set of points in the grid. The task of the logical process is to fill out the points in that set. Our approach throughout this chapter has been to restrict such sets to being vertical stripes, but there is in principle no reason why logical processes may not have different shapes.

Sequential simulation methods can be viewed as employing one single logical process corresponding to the entire grid. The distributed methods we have studied all restrict each vertical stripe to correspond to exactly one physical process. Sequential methods normally fill out the grid in increasing order of time, and essentially this is what their distributed time-stepped and conservative event-driven counterparts do as well. Optimistic event-driven methods also have the overall goal of doing that, but they do it in a rather unsynchronized manner, and occasionally points that have been filled out may have to be erased for later reconsideration.

Although all the methods we have studied adopt such a vertical-stripe approach to filling out the grid, the use of other subdivisions in the distributed case accounts for interesting possibilities. For example, any arrangement other than the one based on vertical stripes requires more than one logical process to simulate the same physical process, however for different time intervals. Whether physical systems exist for which such an arrangement of logical processes is capable of performing efficiently remains largely to be seen.

10.7 Exercises

1. Provide the termination details for the two cases discussed in Section 10.2. Assume, in both cases, that an additional node, \( n_0 \), exists whose function is to detect termination. Provide a solution for each of the following two cases. First, all nodes update the states of their physical processes the same number of times. Second, nodes perform updates until some global convergence criterion is met.
2. Provide an algorithm for the time-stepped simulation of fully concurrent systems when edges are not FIFO.
3. Provide a version of Algorithm A Schedule PC in which the initial propagation of states is selective, depending on the initial orientation.
4. Complete action (10.4) by specifying how to update.
5. Write the algorithm described in Section 10.3.2.
6. Discuss an alternative to the use of anti-event's if G's edges are FIFO.
3. Provide a version of Algorithm A_Schedule_PC in which the initial propagation of states is selective, depending on the initial orientation.

4. Complete action (10.4) by specifying how to update \( T_i \).

5. Write the algorithm described in Section 10.3.2.

6. Discuss an alternative to the use of anti-event's if G's edges are FIFO.

10.8 Bibliographic notes

There are many sources of reference to complement the introductory material presented in Section 10.1. A general introduction to the role played by high-performance computation in the analysis of physical systems is given by Fox and Otto (1986). Readers wanting to concentrate on the simulation of physical systems based on continuous-time may need to acquire a deeper understanding of the mathematics involved (Wylie, 1975; Luenberger, 1979), while those whose interest lies mainly on the study of discrete-time physical systems may wish to check some of the paradigmatic problems in the area, as the firing squad problem (Jiang, 1989) and the chip firing problem (Spencer, 1986). Overviews of various aspects of distributed event-driven simulation can be found in many places, including Misra (1986), Fujimoto (1990a), Fujimoto (1993), and Nicol and Fujimoto (1994). A discussion on the relation between distributed simulation and the notions of knowledge of Section 2.3 has been given by Loucks and Preiss (1990). Alonso, Frutos, and Palacio (1994) provide a comparative study of conservative and optimistic methods, while Lin (1993) discusses the termination of event-driven simulations. Pohllmann (1991) discusses an approach that departs from our taxonomy.

Section 10.2 is entirely based on Barbosa and Lima (1990) and on Barbosa (1991; 1993), where the time-stepped simulation of numerous physical systems is considered. These include cellular automata (von Neumann, 1966; Wolfram, 1986), analog Hopfield neural networks (Hopfield, 1984), neural networks to solve linear systems and linear programming problems (de Carvalho and Barbosa, 1992), binary Hopfield neural networks (Hopfield, 1982), systems under simulated annealing (Kirkpatrick, Gelatt, and Vecchi, 1983; Gafni and Barbosa, 1986; Barbosa and Gafni, 1989a; Barbosa and Boeres, 1990), Markov random fields (Geman and Geman, 1984), Boltzmann machines (Hinton, Sejnowski, and Ackley, 1984), and Bayesian networks (Pearl, 1988; Eizirik, Barbosa, and Mendes, 1993). The pioneering work on conservative methods for distributed simulation (and hence for distributed simulation in general) was done independently by Bryant (1977) and by Chandy and Misra (1979). It is on the work of Chandy and Misra (1979) that Section 10.3.1 is based. Section 10.3.2 is based on the later work by Chandy and Sherman (1989b). In the context of conservative methods, various authors have addressed the question of reducing the number of null messages, as for example De Vries (1990) and Preiss, Loucks, MacIntyre, and Field (1990). Similarly, the effects of the absence of lookaheads (Lin, Lazowska, and Baer, 1990) and of their presence (Preiss and Loucks, 1990) have also been studied. Other studies on conservative methods have been conducted by Chandy and Misra (1981)—where a deadlock-detection strategy is employed in place of prevention, Mehl (1990), Yu, Ghosh, and DeBenedictis (1990), Lin, Lazowska, and Hwang (1992), Nicol (1992), Ayani and Rajaei (1994), Blanchard, Lake, and Turner (1994), Teo and Tay (1994), and Wood and Turner (1994).

Fujimoto (1990b) presents a survey of optimistic methods for distributed simulation, and Preiss, MacIntyre, and Loucks (1992) elaborate on the relation between "optimism" and memory availability in optimistic methods. Section 10.4 on the time warp mechanism is based on Jefferson (1985). For studies on the management of memory during optimistic simulations, the reader is referred to Lin (1994), for a treatment related to optimistic methods in general, and to Lin and Preiss (1991) and Das and Fujimoto (1994), for a treatment within the context of the time warp mechanism. Methods for the periodic computation of GVT have been proposed by many authors, including, more recently, Matern (1993), Srinivasan and Reynolds (1993), and D'Souza, Fan, and Wilsey (1994). Strategies for limiting a method's optimism have also been investigated, and can be looked up, for example, in the works by

Section 10.5 is based on Wedemann, Barbosa, and Donangelo (1995). Details on the physical system described in that section can be found in Aichelin and Bertsch (1985), Bertsch and Gupta (1988), and Bauer, Bertsch, and Schulz (1992).

Material related to our discussion in Section 10.6 is available from Chandy and Sherman (1989a), Bagrodia, Chandy, and Liao (1991), and Jha and Bagrodia (1994).

**Bibliography**


DIMACS (Center for Discrete Mathematics and Theoretical Computer Science) (1990). The first DIMACS international algorithm implementation challenge: general information; problem definitions and specifications; the core experiments. DIMACS, Rutgers University, Piscataway, NJ.


List of Figures

Chapter 1: Message-Passing Systems

Figure 1.1: A graph $G_T$ is shown in part (a). In the graphs of parts (b) through (d), circular nodes are the nodes of $G_T$, while square nodes represent buffers assigned to the corresponding channel in $G_T$. If $r(c) = 1$ for all $c \in \{c_1, c_2, c_3, c_4\}$, then parts (b) through (d) represent three distinct buffer assignments, all of which deadlock-free. Part (b) shows the
strategy of setting \( b(c) = r(c) \) for all \( c \in \{c_1, c_2, c_3\} \). Parts (c) and (d) represent, respectively, the results of the space-optimal and the concurrency-optimal strategies.

Figure 1.2: When task \( u \) migrates from processor \( p \) to processor \( p' \) and \( v \) from \( q \) to \( q' \), a flush\((u, v, p')\) message and a flush-request\((u, v)\) message are sent concurrently, respectively by \( p \) to \( q \) and by \( q \) to \( p \). The flush message gets forwarded by \( q \) to \( q' \), and eventually causes \( q' \) to send \( p' \) a flushed\((u, v, q')\) message.

Chapter 2: Intrinsic Constraints

Figure 2.1: This is the \( 2v \) \((2|T/3| + 1)\)-node ring used in the proof of Theorem 2.4, here shown for \( v = 3 \) and \( T = 3 \). Each of the three portions in the upper half comprising three contiguous nodes each is assigned \( f \)'s arguments according to \( a \). Similar portions in the lower half of the ring follow assignment \( a \).

Figure 2.2: The \( 2v \)-node ring used in the proof of Theorem 2.6 is depicted here for \( v = 5 \). Shown is also the mapping \( \psi: \{1, \ldots, 2v\} \rightarrow \{1, \ldots, 2v\} \), emphasizing the symmetry among the nodes in the ring's upper half and the corresponding nodes in the lower half.

Chapter 3: Models of Computation

Figure 3.1: A precedence graph has \( \Xi \) for node set and the pairs in the partial order \( \prec \) for edges. It is convenient to draw precedence graphs so that events happening at the same node in \( N \) are placed on a horizontal line and positioned on this line, from left to right, in increasing order of the local times at which they happen. In this figure, shown for \( n = 4 \), the "conically"-shaped regions delimited by dashed lines around event \( \xi \) happening at node \( n_3 \) represent \( \{\xi\} \cup \text{Past}(\xi) \) (the one on the left) and \( \{\xi\} \cup \text{Future}(\xi) \) (the one on the right).

Figure 3.2: Part (a) of this figure shows a precedence graph, represented by solid lines, for \( n = 2 \). As \( \prec \) is already transitive, we have \( \prec^+ = \prec \). Members of \( \prec^+ \) are then represented by solid lines, while the dashed lines are used to represent the pairs of concurrent events, which, when added to \( \prec^+ \), yield a total order \( \prec^+ \) consistent with \( \prec \). The same graph is redrawn in part (b) of the figure to emphasize the total order. In this case, system_state \((\xi_2, \xi_3)\) is such that \( n_1 \) is in the state at which it was left by the occurrence of \( \xi_1 \), \( n_2 \) is in the state at which it was left by the occurrence of \( \xi_2 \), and a message sent in connection with \( \xi_2 \) is in transit on the edge from \( n_2 \) to \( n_1 \) to be received in connection with \( \xi_3 \). Because \( \prec^+ \) is consistent with \( \prec \), system_state \((\xi_2, \xi_3)\) is a global state, by our first definition of global states.

Figure 3.3: Parts(a) and (b) show the same precedence graph for \( n = 2 \). Each of the cuts shown establishes a different partition \((\Xi_1, \Xi_2)\) of \( \Xi \). The cut in part (a) has no edge leading from an event in \( \Xi_2 \) to an event in \( \Xi_1 \), and then system_state \((\Xi_1, \Xi_2)\) is a global state, by our second definition. In this global state, \( n_1 \) is in its initial state, \( n_2 \) is in the state at which it was left by the occurrence of \( \xi_2 \), and a message is in transit on the edge from \( n_2 \) to \( n_1 \), sent in connection with \( \xi_2 \), and to be received in connection with \( \xi_3 \). The cut in part (b), on the other hand, has an edge leading from \( \xi_2 \in \Xi_2 \) to \( \xi_3 \in \Xi_1 \), so system_state\((\Xi_1, \Xi_2)\) cannot be a global state.

Chapter 4: Basic Algorithms

Figure 4.1: During an execution of Algorithm A_PIF, the variables parent \( i \), for all nodes \( n_i \) are set so that a spanning tree is created on \( G \). This spanning tree is rooted at \( n_1 \), and its leaves correspond to nodes from which no other node received \( \text{in} \) \( f \) for the first time. In this figure, a directed edge is drawn from \( n_i \) to \( n_j \) to indicate that parent \( i \neq n_j \).

Chapter 6: Stable Properties

Figure 6.1: Edges in the precedence graph fragment shown in part (a) are drawn as either solid lines or dashed lines. Solid lines represent comp_msg's, dashed lines represent ack's, and the remaining edges of the precedence graph are omitted. In this case, system_state \((\Xi_1, \Xi_2)\) is clearly a global state, and is such that every node that ever sent a comp_msg during the diffusing computation (i.e., \( n_1 \) and \( n_2 \)) is in the state that immediately precedes the reception of the last ack. In part (b), the spanning tree formed by the variables parent for each node \( n_i \) in this global state is shown with directed edges that point from \( n_i \) to \( n_j \) to indicate that parent \( i \neq n_j \). In this case, the tree has \( n_1 \) for root and its single leaf is \( n_3 \).
Chapter 8: Resource Sharing

Figure 8.1: A period of five orientations results from the edge-reversal mechanism started at the orientation shown in the upper left corner of the figure, which is outside the period. In this period, every node becomes a sink twice.

Chapter 9: Program Debugging

Figure 9.1: In this figure, the solid segment in a process's horizontal line indicates the time interval during which the corresponding local predicate is true. The two cuts shown clearly correspond to global states, in fact earliest global states in which the disjunctive predicate holds.

Figure 9.2: The tiny solid segment in a process's horizontal line indicates the local time to which the corresponding local unconditional breakpoint has been set. Clearly, the settings in this figure are erroneous, as the cut (shown as a dashed line) that goes through them does not correspond to a global state.

Figure 9.3: Following the same conventions as in Figure 9.2, here a situation is depicted in which only one node participates in the unconditional breakpoint (node n). Depending on how the corresponding local unconditional breakpoint is placed with respect to the reception of the message by p, the other processes appear in the resulting earliest global state differently, as shown in parts (a) and (b).

Figure 9.4: The conventions employed in this figure are the same as those of Figure 9.1, and the situation depicted is quite akin to that of Figure 9.3. Specifically, the earliest global state at which the conjunctive predicate holds depends on when n's local predicate becomes true with respect to the reception of the message by p, as shown in parts (a) and (b).