Chapter 12

DISTRIBUTED DBMS RELIABILITY

The purpose of this chapter is to discuss the reliability features of a distributed DBMS. From Chapter 10 the reader will recall that the reliability of a distributed DBMS refers to the atomicity and durability properties of transactions. Two specific aspects of reliability protocols that need to be discussed in relation to these properties are the commit and the recovery protocols. In that sense, in this chapter we relax the major assumption of Chapter 11 that the underlying distributed system is fully reliable and does not experience any hardware or software failures. Furthermore, the commit protocols discussed in this chapter constitute the support provided by the distributed DBMS for the execution of commit commands which we have placed in transactions.

It is possible to discuss database reliability in isolation. However, the distributed DBMS is only one component of a distributed computing system. Its reliability is strongly dependent on the reliability of the hardware and software components that make up the distributed environment. Therefore, our discussion in this chapter starts with a general presentation of the reliability issues in distributed computing systems, and then focuses on the reliability aspects of distributed databases.

12.1 RELIABILITY CONCEPTS AND MEASURES

Too often, the terms *reliability* and *availability* are used loosely in literature. Even among the researchers in the area of reliable computer systems, there is no consensus on the definitions of these terms. In this section we give precise definitions of a number of concepts that are fundamental to an understanding and study of reliable systems. Our definitions follow those of [Anderson and Lee, 1985] and [Randell et al., 1978]. Nevertheless, we indicate where these definitions might differ from other usage of the terms.

12.1.1 System, State, and Failure

In the context of reliability, *system* refers to a mechanism that consists of a collection of components and interacts with its environment by responding to stimuli from the environment with a recognizable pattern of behavior (Figure 12.1). Each component of a system is itself a system, commonly called a *subsystem*. The environment of a component is the system of which it is a part. The way the components of a system are put together is called the *design* of the system. For example, Figure 4.7 depicts a distributed DBMS with all its components and the design of the DBMS, which indicates the interactions between these components. Note in Figure 4.7 that we have depicted only the software components. The only hardware component that is shown is the disk drive. It is, however, possible and even necessary to define the hardware components if one has to consider the reliability of the entire DBMS.

There are a number of ways of modeling the interaction between the software and the hardware in a computer system. One possible modeling method is to treat the program text as the design of an abstract system whose components are the hardware and software objects that are manipulated during the execution of the program. Another modeling alternative is to specify each software and hardware component explicitly, and let the design represent the way these components interact with one another (similar to Figure 4.7).

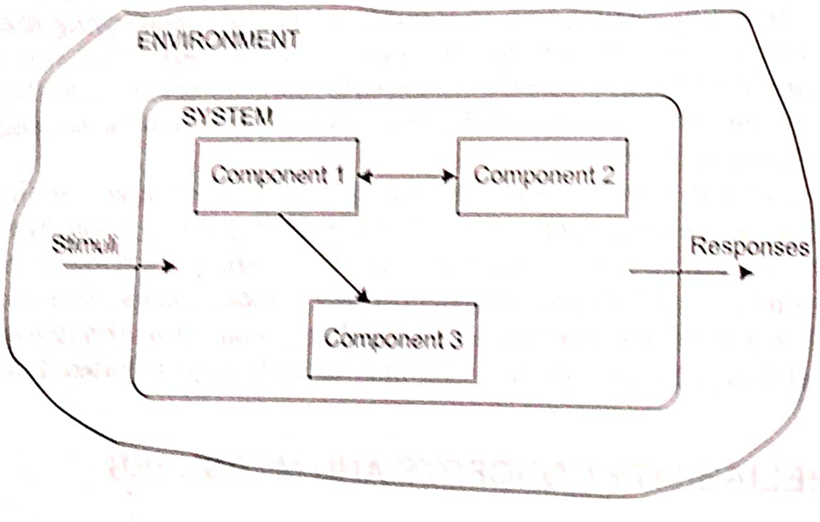


Figure 12.1. Schematic of a System

An *external state* of a system can be defined as the response that a system gives to an external stimulus. It is therefore possible to talk about a system changing external states according to repeated stimuli from the environment. We can define the *internal state* of the system similarly. It is convenient to define the internal state as the union of the external states of the components that make up the system. Again, the system changes its internal state in response to stimuli from the environment.

The behavior of the system in providing response to all the possible stimuli from the environment needs to be laid out in an authoritative *specification* of its behavior. The specification indicates the valid behavior of each system state. Such a specification is not only necessary for a successful system design but is also essential to define the following reliability concepts.

Any deviation of a system from the behavior described in the specification is considered a *failure*. For example, in a distributed transaction manager the specification would state that only serializable schedules for the execution of concurrent transactions should be generated. If the transaction manager generates a nonserializable schedule, we say that it has failed.

Each failure obviously needs to be traced back to its cause. Failures in a system can be attributed to deficiencies either in the components that make it up, or in the design, that is, how these components are put together. Each state that a reliable system goes through is valid in the sense that the state fully meets its specification. However, in an unreliable system, it is possible that the system may get to an internal state that may not obey its specification. Further transitions from this state would eventually cause a system failure. Such internal states are called *erroneous states*; the part of the state that is incorrect is called an *error* in the system. Any error in the internal states of the components of a system or in the design of a system is called a *fault* in the system. Thus a fault causes an error that results in a system failure (Figure 12.2).

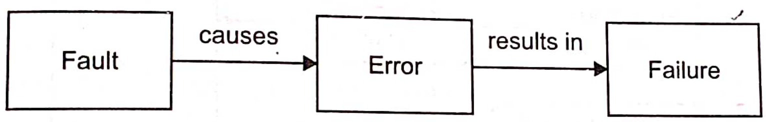


Figure 12.2. Chain of Events Leading to System Failure

We differentiate between errors (or faults and failures) that are permanent and those that are not permanent. Permanence can apply to a failure, a fault, or an error, although we typically use the term with respect to faults. A *permanent fault*, also commonly called a *hard fault*, is one that reflects an irreversible change in the behavior of the system. Permanent faults cause permanent errors that result in permanent failures. The characteristics of these failures is that recovery from them requires intervention to "repair" the fault. Systems also experience *intermittent* and *transient faults*. In the literature, these two are typically not differentiated; they are jointly called *soft faults*. The dividing line in this differentiation is the repairability of the system that has experienced the fault [Siewiorek and Swarz, 1982]. An intermittent fault refers to a fault that demonstrates itself occasionally due to unstable hardware or varying hardware or software states. A typical example is the faults that systems may demonstrate when the load becomes too heavy. On the other hand, a transient fault describes a fault that results from temporary environmental conditions. A transient fault might occur, for example, due to a sudden increase in the room temperature. The transient fault is therefore the result of environmental conditions that may be impossible to repair. An intermittent fault, on the other hand, can be repaired since the fault can be traced to a component of the system.

Remember that we have also indicated that system failures can be due to design faults. Design faults together with unstable hardware cause intermittent errors which result in system failure. A final source of system failure that may not be attributable to a component fault or a design fault is operator mistakes. These are the sources of a significant number of errors as the statistics included further in this section demonstrate. The relationship between various types of faults and failures is depicted in Figure 12.3.

As indicated previously, there are a number of different definitions of the fundamental reliability terms that we discussed. For example, the terms *fault* and *failure* are used interchangeably in the fault tolerance literature, whereas the terms *failure* and *error* are used interchangeably in the coding theory literature [Elkind, 1982]. Alternatively, a *fault* may be defined as "an erroneous state of hardware or software resulting from failures of components" and an *error* as "a manifestation of a fault within a program or data structure." *Failure* is defined as a "physical change in hardware" [Avizienis, 1977]. These definitions are specific to hardware systems and represent a rather different view as to what causes what. According to these definitions, a failure causes a fault that results in an error.

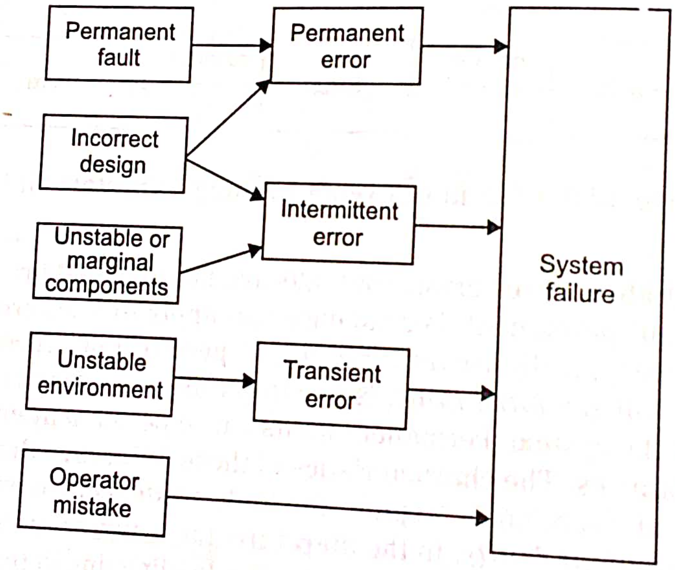


Figure 12.3. Sources of System Failure (Reprinted with permission from *The Theory and Practice of Reliable System Design*, by Daniel P. Siewiorek and Robert S. Swarz, copyright ©Digital Press/Digital Equipment Corporation, 12 Crosby Drive, Bedford, MA 01730.)

12.1.2 Reliability and Availability

Reliability refers to the probability that the system under consideration does not experience any failures in a given time interval. It is typically used to describe systems that cannot be repaired (as in space-based computers), or where the operation of the system is so critical that no downtime for repair can be tolerated.

Formally, the reliability of a system, $R(t)$, is defined as the following conditional probability:

$$R(t) = \Pr\{0 \text{ failures in time } [0, t] \mid \text{no failures at } t = 0\}$$

Reliability theory, as it applies to hardware systems, has been developed significantly. Let us therefore illustrate the foregoing formula for hardware components, for which it is customary to assume that failures follow a Poisson distribution. In this case

$$R(t) = \Pr\{0 \text{ failures in time } [0, t]\}$$

Under the same assumptions it is possible to derive that

$$\Pr\{k \text{ failures in time } [0, t]\} = \frac{e^{-m(t)}[m(t)]^k}{k!}$$

where $m(t) = \int_0^t z(x)\,dx$. Here $z(t)$ is known as the *hazard function,* which gives the time-dependent failure rate of the specific hardware component under consideration. The probability distribution for $z(t)$ may be different for different electronic components.

The expected (mean) number of failures in time $[0, t]$ can then be computed as

$$E[k] = \sum_{k=0}^{\infty} k \frac{e^{-m(t)}[m(t)]^k}{k!} = m(t)$$

and the variance as

$$Var[k] = E[k^2] - (E[k])^2 = m(t)$$

Given these values, $R(t)$ can be written as

$$R(t) = e^{-m(t)}$$

Note that the reliability equation above is written for one component of the system. For a system that consists of n nonredundant components (i.e., they all have to function properly for the system to work) whose failures are independent, the overall system reliability can be written as

$$R_{sys}(t) = \Pi_{i=1}^{n} R_i(t)$$

Availability, $A(t)$, refers to the probability that the system is operational according to its specification at a given point in time t. A number of failures may have occurred prior to time t, but if they have all been repaired, the system is available at time t. It is apparent that availability refers to systems that can be repaired.

If one looks at the limit of availability as time goes to infinity, it refers to the expected percentage of time that the system under consideration is available to perform useful computations. Availability can be used as some measure of "goodness" for those systems that can be repaired and which can be out of service for short periods of time during repair. Reliability and availability of a system are considered to be contradictory objectives [Siewiorek and Swarz, 1982]. It is usually accepted that it is easier to develop highly available systems as opposed to highly reliable systems.

If we assume that failures follow a Poisson distribution with a failure rate λ, and that repair time is exponential with a mean repair time of $1/\mu$, the steady-state availability of a system can be written as

$$A = \frac{\mu}{\lambda + \mu}$$

12.1.3 Mean Time between Failures/Mean Time to Repair

Calculation of the reliability and the availability functions is quite tedious. It is therefore customary to use two single-parameter metrics to model the behavior of systems. The two measures used are *mean time between failures* (MTBF) and *mean time to repair* (MTTR). MTBF is the expected time between subsequent failures in a system with repair.[1] MTBF can be calculated either from empirical data or from the reliability function as

$$\text{MTBF} = \int_0^\infty R(t)\,dt$$

Since $R(t)$ is related to the system failure rate, there is a direct relationship between MTBF and the failure rate of a system. MTTR is the expected time to repair a failed system. It is related to the repair rate as MTBF is related to the failure rate. Using these two metrics, the steady-state availability of a system with exponential failure and repair rates can be specified as

$$A = \frac{\text{MTBF}}{\text{MTBF} + \text{MTTR}}$$

12.2 FAILURES AND FAULT TOLERANCE IN DISTRIBUTED SYSTEMS

In this section we consider the reasons for failures in distributed systems as well as the basic fault-tolerance techniques that are used to cope with them. This discussion is based on empirical statistics and is not meant to be complete and exhaustive. It is aimed only at providing a general framework for the distributed database reliability issues.

12.2.1 Reasons for Failures

Let us first take a look at soft and hard failures. Soft failures make up more than

[1]A distinction is sometimes made between MTBF and MTTF (mean time to fail). MTTF is defined as the expected time of the first system failure given a successful startup at time 0. MTBF is then defined only for systems that can be repaired. An approximation for MTBF is given as MTBF = MTTF + MTTR [McConnel and Siewiorek, 1982]. We do not make this distinction in this book.

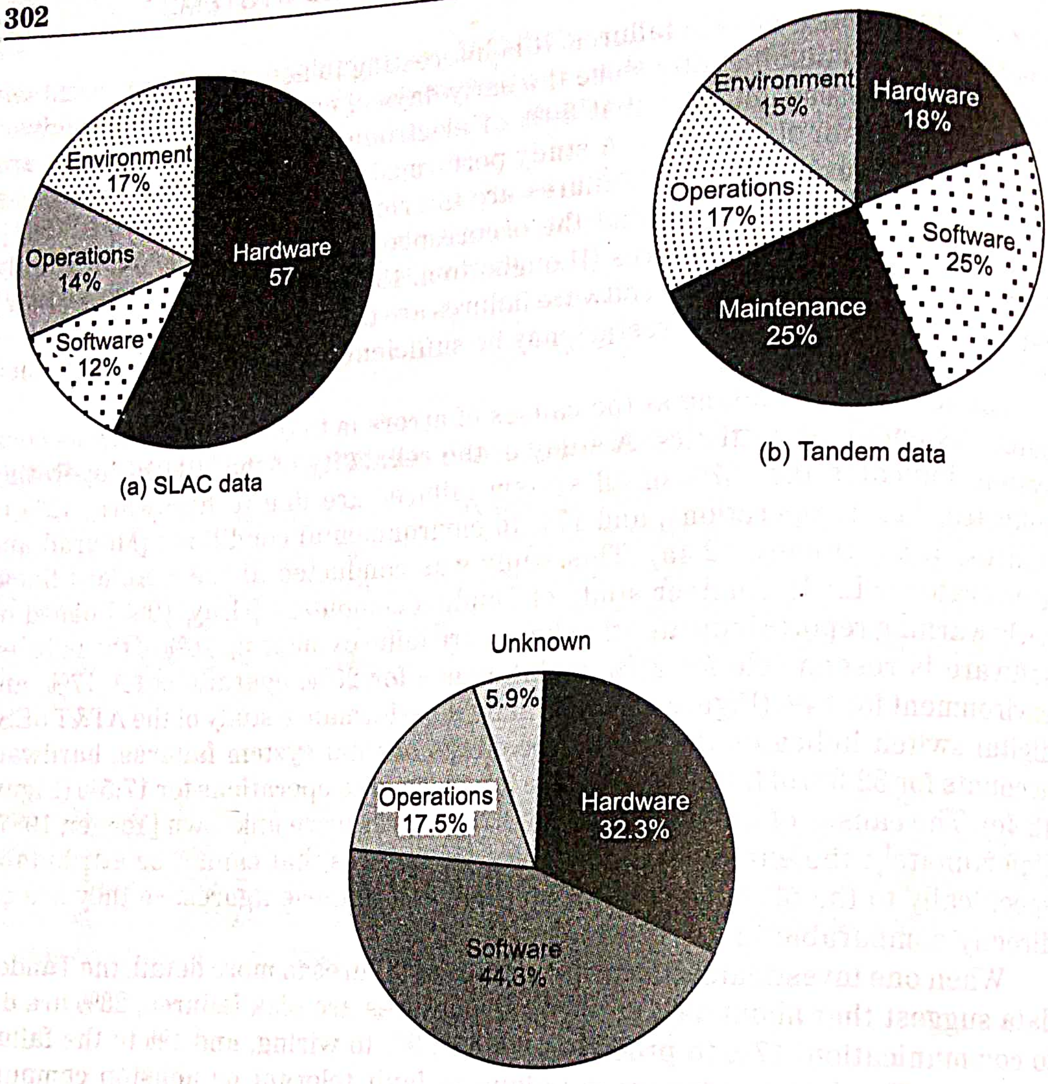
90% of all hardware system failures. It is interesting to note that this percentage has not changed significantly since the early days of computing. A 1967 study of the U.S. Air Force indicates that 80% of electronic failures in computers are intermittent [Roth et al., 1967]. A study performed by IBM during the same year concludes that over 90% of all failures are intermittent [Ball and Hardie, 1967]. More recent studies indicate that the occurrence of soft failures is significantly higher than that of hard failures ([Longbottom, 1980], [Gray, 1987]). Gray [1987] also mentions that most of the software failures are transient—and therefore soft—suggesting that a dump and restart may be sufficient to recover without any need to "repair" the software.

Another way of looking at the causes of errors is to investigate various computer system error statistics. A study of the reliability of the IBM/XA operating system indicates that 57% of all system failures are due to hardware, 12% to software, 14% to operations, and 17% to environmental conditions [Mourad and Andres, 1985] (Figure 12.4a). This study was conducted at the Stanford linear accelerator (SLAG). Another study of Tandem computers [Gray, 1985] based on early warning reports indicates that hardware failures make up 18% of the failures, software is responsible for 25%, maintenance for 25%, operations for 17%, and environment for 14% (Figure 12.4b). Finally, a performance study of the AT&T 5ESS digital switch indicates that as a percentage of total system failures, hardware accounts for 32.3% of failures, software for 44.3%, and operations for 17.5% (Figure 12.4c). The causes of the remaining 5.9% of failures were unknown [Yeager, 1987]. Unfortunately, the environmental causes for failures that cannot be attributable specifically to the 5ESS switch are not included in these figures, so they are not directly comparable to the previous two.

When one investigates hardware causes of failures in more detail, the Tandem data suggest that about 49% of hardware failures are disk failures, 23% are due to communication, 17% to processor failure, 9% to wiring, and 1% to the failure of spares. The latter category is unique to fault tolerant or nonstop computer systems such as Tandem since they use spare modules to achieve fault tolerance. In the case of the 5ESS switch, 32.5% of the hardware failures are due to disk failures, 9.3% are due to electromagnetic interference due to insufficient isolation of wires, and the remaining 58.2% are due to defective components, circuit packs, or cables.

Software failures are more difficult to discuss because there is no agreement on a classification scheme. The Tandem data again suggests that software failures are by far the dominant causes. These are due to communication and database failures, followed by operating system failures, which are then followed by failures in the application code and in the transaction management software.

Software failures are typically caused by "bugs" in the code. The estimates for the number of bugs in software vary considerably. Figures such as 0.25 bug per 1000 instructions to 10 bugs per 1000 instructions have been reported. As stated before, most of the software failures are soft failures. The statistics for software failure are comparable to those we have previously reported on hardware



(c) 5ESS switch data

Figure 12.4. Reasons for Computer System Failures

failures. The fundamental reason for the dominance of soft failures in software is the significant amount of design review and code inspection that a typical software project goes through before it gets to the testing stage. Furthermore, most commercial software goes through extensive alpha and beta testing before being released for field use.

An interesting classification for software bugs which derives from physics can also be provided [Gray, 1985]. The classification divides software bugs into Heisenbugs and Bohrbugs. Bohrbugs cause hard faults; they are solid like a Bohr atom and will continue to cause faults at retry. Heisenbugs, on the other hand, behave according to the Heisenberg uncertainty principle in physics, which states that it is not possible to measure both the position and the velocity of an electron accurately and simultaneously. The measurement procedure for one of these will cause a change in the other one. The implication of this principle for software bugs is that the testing process to find a bug may cause sufficient perturbation for the bug to disappear. Thus Heisenbugs cause soft faults and are more challenging to detect. They are sensitive to time of execution, system load, and other similar factors and the system may work on retry.

12.2.2 Basic Fault Tolerance Approaches and Techniques

The two fundamental approaches to constructing a reliable system are fault tolerance and fault prevention. *Fault tolerance* refers to a system design approach which recognizes that faults will occur; it tries to build mechanisms into the system so that the faults can be detected and removed or compensated for before they can result in a system failure. *Fault prevention* techniques, on the other hand, aim at ensuring that the implemented system will not contain any faults. Fault prevention has two aspects. The first is *fault avoidance* which refers to the techniques used to make sure that faults are not introduced into the system. These techniques involve detailed design methodologies (such as design walkthroughs, design inspections, etc.) and quality control. The second aspect of fault prevention is *fault removal*, which refers to the techniques that are employed to detect any faults that might have remained in the system despite the application of fault avoidance and removes these faults. Typical techniques that are used in this area are extensive testing and validation procedures. Note that these fault removal techniques apply during system implementation prior to field use of the system.

The terms *fault prevention* and *fault avoidance* are used interchangably. Another common name for these approaches is *fault intolerance* [Avizienis, 1976] since they cannot withstand faults that show up during system use. These techniques concentrate on designing systems using high-reliability components and refinement of the packaging techniques followed by extensive testing. Such measures are expected to reduce the occurrence of system failures to a minimum so that the features may then be handled by manual maintenance. Unfortunately, there are a number of environments where manual maintenance and repair is impossible, or the downtime needed for repair would be intolerable. In those environments fault tolerant system designs are the preferred alternative.

It is possible to talk about a third approach to constructing reliable systems. This is *fault detection* [Myers, 1976]. Specification of fault detection as a separate approach may be challenged since detection has to be included in any fault tolerance technique. However, if fault detection techniques are not coupled with fault tolerance features, they issue a warning when a failure occurs but do not provide any means of tolerating the failure. Therefore, it might be appropriate to separate fault detection from strictly fault tolerant approaches.

It is important to note at this point that system failures may be *latent*. A latent failure is one that is detected some time after its occurrence. This period is called *error latency*, and the average error latency time over a number of identical systems is called *mean time to detect* (MTTD). Figure 12.5 depicts the relationship of various reliability measures with the actual occurrences of faults.

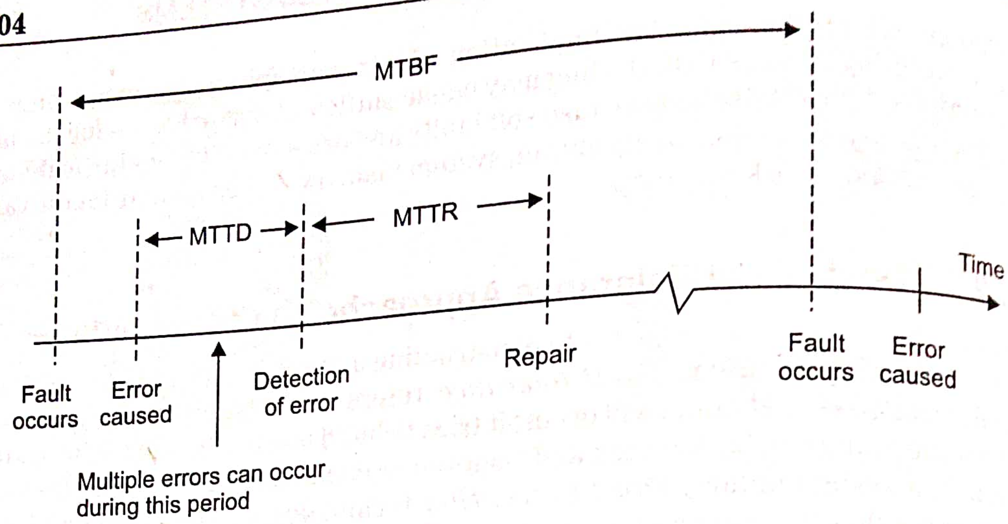


Figure 12.5. Occurrence of Events over Time

The fundamental principle employed in all fault tolerant system designs is that of providing *redundancy* in system components. Redundant components enable the effects of a faulty component to be compensated for. However, redundancy is not sufficient for fault tolerance. The additional and complementary fault tolerance principle is the *modularization* of the design. Each component of the system is implemented as a module with well-defined input and output interfaces with the other components. Modularization enables the isolation of faults within one component. It is therefore an important technique in both hardware and software systems.

Process pairs require communication between processes. An argument may be made for uniprocessor systems that, for the sake of improved performance, the communication between processes may be implemented by means of shared memory. However, when designing a reliable software environment, it is important to implement an operating system that uses a message-based interprocess communication mechanism. Such an approach contributes to fault isolation since each process executes in its own address space, and an error that one of them might cause will not propagate to the other processes.

Another important related concept is session-oriented communication between processes. Session-oriented communication delegates the responsibility of detecting and handling lost or duplicate messages to the message server of the operating system rather than to the application program. This not only facilitates a simpler application development environment but also enables the operating system to provide a reliable execution environment to application processes.

12.3 FAILURES IN DISTRIBUTED DBMS

Designing a reliable system that can recover from failures requires identifying the types of failures with which the system has to deal. In Section 12.2 we reviewed the major reasons for failures in distributed computer systems. That discussion, together with the termination conditions of transactions, covered in Chapter 10, indicates that a database recovery manager has to deal with four types of failures: transaction failures (aborts), site (system) failures, media (disk) failures, and communication line failures.

12.3.1 Transaction Failures

Transactions can fail for a number of reasons. Failure can be due to an error in the transaction caused by incorrect input data (e.g., Example 10.3) as well as the detection of a present or potential deadlock. Furthermore, some concurrency control algorithms do not permit a transaction to proceed or even to wait if the data that they attempt to access are currently being accessed by another transaction. This might also be considered a failure. The usual approach to take in cases of transaction failure is to *abort* the transaction, thus resetting the database to its state prior to the start of this transaction.

The frequency of transaction failures is not easy to measure. It is indicated that in System R, 3% of the transactions abort abnormally [Gray et al., 1981]. In general, it can be stated that (1) within a single application, the ratio of transactions that abort themselves is rather constant, being a function of the incorrect data, the available semantic data control features, and so on; and (2) the number of transaction aborts by the DBMS due to concurrency control considerations (mainly deadlocks) is dependent on the level of concurrency (i.e., number of concurrent transactions), the interference of the concurrent applications, the granularity of locks, and so on [Härder and Reuter, 1983].

12.3.2 Site (System) Failures

The reasons for system failure constituted the bulk of our discussion in Section 12.2. In short, it can be traced back to a hardware failure (processor, main memory, power supply, etc.) or to a software failure (bug in the operating system or in the DBMS code). The important point from the perspective of this discussion is that a system failure is always assumed to result in the loss of main memory contents. Therefore, any part of the database that was in main memory buffers is lost as a result of a system failure. However, the database that is stored in secondary storage is assumed to be safe and correct. In distributed database terminology, system failures are typically referred to as *site failures*, since they result in the failed site being unreachable from other sites in the distributed system.

We typically differentiate between partial and total failures in a distributed system. *Total failure* refers to the simultaneous failure of all sites in the distributed system; *partial failure* indicates the failure of only some sites while the others remain operational. As indicated in Chapter 1, it is this aspect of distributed systems that makes them more available.

12.3.3 Media Failures

Media failure refers to the failures of the secondary storage devices that store the database. Such failures may be due to operating system errors, as well as to hardware faults such as head crashes or controller failures. The important point from the perspective of DBMS reliability is that all or part of the database that is on the secondary storage is considered to be destroyed and inaccessible.

Duplexing of disk storage and maintaining archival copies of the database are common techniques that deal with this sort of catastrophic problem. Even though the data of Section 12.2 suggest that disk failures are quite common, the techniques described above enable us to assume that media failures do not have an impact more often than once or twice a year [Härder and Reuter, 1983].

Media failures are frequently treated as problems local to one site and therefore not specifically addressed in the reliability mechanisms of distributed DBMSs. We consider techniques for dealing with them in Section 12.4.5 under local recovery management. We then turn our attention to site failures when we consider distributed recovery functions.

12.3.4 Communication Failures

The three types of failures described above are common to both centralized and distributed DBMSs. Communication failures, however, are unique to the distributed case. There are a number of types of communication failures. The most common ones are the errors in the messages, improperly ordered messages, lost (or undeliverable) messages, and line failures. As discussed in Chapter 3, the first two errors are the responsibility of the computer network (specifically, the communication subnet, which consists of the physical, data link, and network layers of the ISO/OSI architectural model); we will not consider them further. Therefore, in our discussions of distributed DBMS reliability, we expect the underlying computer network hardware and software to ensure that two messages sent from a process at some originating site to another process at some destination site are delivered without error and in the order in which they were sent.

Lost or undeliverable messages are typically the consequence of communication line failures or (destination) site failures. If a communication line fails, in addition to losing the message(s) in transit, it may also divide the network into two or more disjoint groups. This is called *network partitioning*. If the network is partitioned, the sites in each partition may continue to operate. In this case, executing transactions that access data stored in multiple partitions becomes a major issue. Maintaining the mutual consistency of the database is a significant problem, especially if the database is replicated across these partitions.

Network partitions point to a unique aspect of failures in distributed computer systems. In centralized systems the system state can be characterized as all-or-nothing: either the system is operational or it is not. Thus the failures are complete: when one occurs, the entire system becomes nonoperational. Obviously, this is not true in distributed systems. As we indicated a number of times before, this is their potential strength. However, it also makes the transaction management algorithms more difficult to design.

If messages cannot be delivered, we will assume that the network does nothing about it. It will not buffer it for delivery to the destination when the service is reestablished and will not inform the sender process that the message cannot be delivered. In short, the message will simply be lost. We make this assumption because it represents the least expectation from the network and places the responsibility of dealing with these failures to the distributed DBMS.

One result of this assumption is that the responsibility of detecting that a message is undeliverable is left to the application program (in this case the distributed DBMS). The detection will be facilitated by the use of timers and a timeout mechanism that keeps track of how long it has been since the sender site has not received a confirmation from the destination site about the receipt of a message. This timeout interval needs to be set to a value greater than that of the maximum round-trip propagation delay of a message in the network. The term for the failure of the communication network to deliver messages and the confirmations within this period is *performance failure*. It needs to be handled within the reliability protocols for distributed DBMSs.

12.4 LOCAL RELIABILITY PROTOCOLS

In this section we discuss the functions performed by the local recovery manager (LRM) that exists at each site. These functions maintain the atomicity and durability properties of local transactions. They relate to the execution of the commands that are passed to the LRM, which are **begin_transaction**, **read**, **write**, **commit**, and **abort**. Later in this section we introduce a new command into the LRM's repertoire that initiates recovery actions after a failure. Note that in this section we discuss the execution of these commands in a centralized environment. The complications introduced in distributed databases are addressed in the upcoming sections.

12.4.1 Architectural Considerations

It is again time to use our architectural model (Figures 4.7 and 10.4) and discuss the specific interface between the LRM and the database buffer manager (BM). Remember that all accesses to the database are via the database buffer manager. The detailed discussion of the algorithms that the buffer manager implements is beyond the scope of this book; we provide a summary later in this subsection. Even without these details, we can still specify the interface and its function, as depicted in Figure 12.6.[2]

In this discussion we assume that the database is stored permanently on secondary storage, which in this context is called the *stable storage* [Lampson and Sturgis, 1976]. The stability of this storage medium is due to its robustness to failures. A stable storage device would experience considerably less-frequent failures than would a nonstable storage device. In today's technology, stable storage is typically implemented by means of duplexed magnetic disks which store duplicate copies of data that are always kept mutually consistent. We call the

[2]This architectural model is similar to that given in [Härder and Reuter, 1983] and [Bernstein et al., 1987].

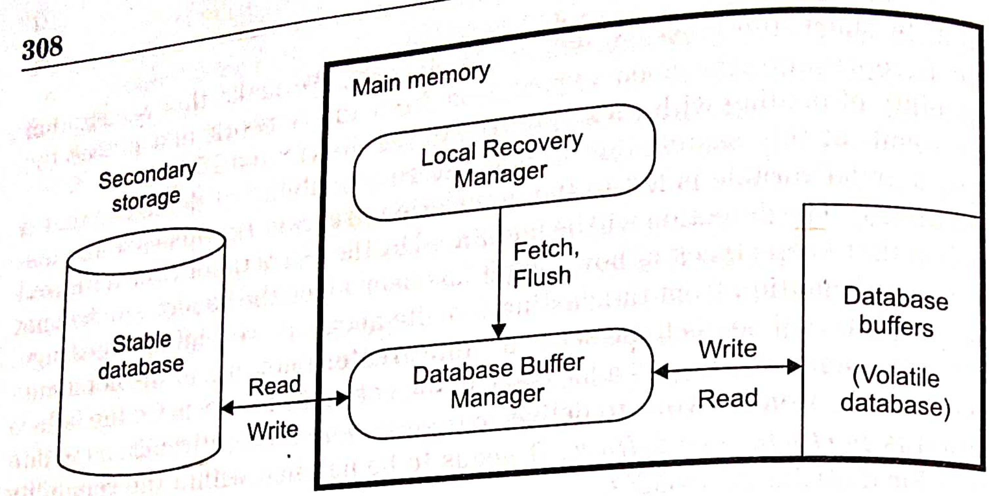


Figure 12.6. Interface Between the Local Recovery Manager and the Buffer Manager

version of the database that is kept on stable storage the *stable database*. The unit of storage and access of the stable database is typically a *page*.

The database buffer manager keeps some of the recently accessed data in main memory buffers. This is done to enhance access performance. Typically, the buffer is divided into pages that are of the same size as the stable database pages. The part of the database that is in the database buffer is called the *volatile database*. It is important to note that the LRM executes the operations on behalf of a transaction only on the volatile database, which, at a later time, is written back to the stable database.

When the LRM wants to read a page of data[3] on behalf of a transaction—strictly speaking, on behalf of some operation of a transaction—it issues a **fetch** command, indicating the page that it wants to read. The buffer manager checks to see if that page is already in the buffer (due to a previous fetch command from another transaction) and if so, makes it available for that transaction; if not, it reads the page from the stable database into an empty database buffer. If no empty buffers exist, it selects one of the buffer pages to write back to stable storage and reads the requested stable database page into that buffer. There are a number of different algorithms by which the buffer manager may choose the buffer page to be replaced; these are discussed in standard database textbooks.

The buffer manager also provides the interface by which the LRM can actually force it to write back some of the buffer pages. This can be accomplished by means of the **flush** command, which specifies the buffer pages that the LRM wants to be written back. We should indicate that different LRM implementations may or may not use this forced writing. This issue is discussed further in subsequent sections.

As its interface suggests, the buffer manager acts as a conduit for all access to the database via the buffers that it manages. It provides this function by fulfilling three tasks:

[3]The LRM's unit of access may be in blocks which have sizes different from a page. However, for simplicity, we assume that the unit of access is the same.

1. *Searching* the buffer pool for a given page;
2. If it is not found in the buffer, *allocating* a free buffer page and *loading* the buffer page with a data page that is brought in from secondary storage;
3. If no free buffer pages are available, choosing a buffer page for *replacement*.

Searching is quite straightforward. Typically, the buffer pages are shared among the transactions that execute against the database, so search is global.

Allocation of buffer pages is typically done dynamically. This means that the allocation of buffer pages to processes is performed as processes execute. The buffer manager tries to calculate the number of buffer pages needed to run the process efficiently and attempts to allocate that number of pages. The best known dynamic allocation method is the *working-set algorithm* [Denning, 1968], [Denning, 1980].

A second aspect of allocation is fetching data pages. The most common technique is *demand paging*, where data pages are brought into the buffer as they are referenced. However, a number of operating systems prefetch a group of data pages that are in close physical proximity to the data page referenced. Buffer managers choose this route if they detect sequential access to a file.

In replacing buffer pages, the best known technique is the least recently used (LRU) algorithm that attempts to determine the *logical reference strings* [Effelsberg and Härder, 1984] of processes to buffer pages and to replace the page that has not been referenced for an extended period. The anticipation here is that if a buffer page has not been referenced for a long time, it probably will not be referenced in the near future.

The techniques discussed above are the most common. Other alternatives are discussed in [Effelsberg and Härder, 1984].

Clearly, these functions are similar to those performed by operating system (OS) buffer managers. However, quite frequently, DBMSs bypass OS buffer managers and manage disks and main memory buffers themselves due to a number of problems (see, e.g., [Stonebraker, 1981]). These problems can be listed as follows:

1. In dynamically allocating buffer pages, the OS buffer manager assumes that page references of processes are random and varying over time. Thus the number of buffer pages that are allocated to a process can grow or shrink during its execution. However, it is possible for a DBMS to determine, with significant regularity, its reference pattern to data pages, which allows a more precise calculation of the buffer requirements [Chou, 1985]. In database literature, a number of different algorithms have been proposed such as the hot set model [Sacco and Schkolnick, 1986] and DBMIN [Chou, 1985], [Chou and DeWitt, 1986].

2. Along the same lines, it may not be useful for the OS buffer manager to prefetch data pages when it detects sequential access. Since the DBMS knows its reference pattern, it is in a significantly better position to make that judgment. The reference string is known by the DBMS either as a

result of query compilation (e.g., DB2 [Date, 1984]) or during execution "at (or very shortly after) the beginning of its examination of a block" [Stonebraker, 1981] (e.g., INGRES [Stonebraker et al., 1976]).

3. The LRU replacement algorithm, which performs nicely for general-purpose computing, fails to perform adequately in a number of cases. For example, in a nested-loop join, the buffer page that contains one data page of the outer relation is the least recently used if a replacement page is needed to bring in the last data page of the inner relation. However, if the outer relation buffer page is replaced, it will be read in again immediately, causing a buffer fault. Every page reference from then on will result in a buffer fault.

4. The LRU replacement algorithm writes the buffer pages to secondary storage when it needs a new buffer page. This is called *delayed writing*. However, when the DBMS commits a transaction, it expects the log pages about the data items that the transaction has updated to be stored immediately on stable storage (i.e., disk). If the buffer manager keeps these buffer pages around, it may not be possible to guarantee the durability of transactions when failures occur unless certain protocols are followed (we discuss these later). The OS buffer managers are now aware of these protocols.

5. OS buffers are generally mapped to virtual memory, so database buffers would also be in the virtual address space. Therefore, accessing the buffers may not only cause buffer faults to bring in data pages, but may also cause memory faults since the buffer page may not be mapped to a main memory page frame and there may not be any free frames. This phenomenon is commonly called the *double paging problem* [Chou and DeWitt, 1986].

The consequence of these problems is that DBMSs usually manage their own buffer space and treat disks as raw storage devices. This is a duplication of OS services, but is unavoidable with the current DBMS and OS architectures since these two systems do not communicate and collaborate with each other very effectively.

12.4.2 Recovery Information

In this section we assume that only system failures occur. We defer the discussion of techniques for recovering from media failures until later. Since we are dealing with centralized database recovery, communication failures are not applicable.

When a system failure occurs, the volatile database is lost. Therefore, the DBMS has to maintain some information about its state at the time of the failure in order to be able to bring the database to the state that it was in when the failure occurred. We call this information the *recovery information*.

The recovery information that the system maintains is dependent on the method of executing updates. Two possibilities are in-place updating and out-of-place updating. *In-place updating* physically changes the value of the data item in the stable database. As a result, the previous values are lost. *Out-of-place updating*, on the other hand, does not change the value of the data item in the stable database but maintains the new value separately. Of course, periodically, these updated values have to be integrated into the stable database. We should note that the reliability issues are somewhat simpler if in-place updating is not used. However, most DBMSs use it due to its improved performance.

In-Place Update Recovery Information

Since in-place updates cause previous values of the affected data items to be lost, it is necessary to keep enough information about the database state changes to facilitate the recovery of the database to a consistent state following a failure. This information is typically maintained in a database log. Thus each update transaction not only changes the database but is also recorded in the *database log* (Figure 12.7). Before we discuss the contents of the log, let us first see why this information is necessary for the database to recover to a consistent state.

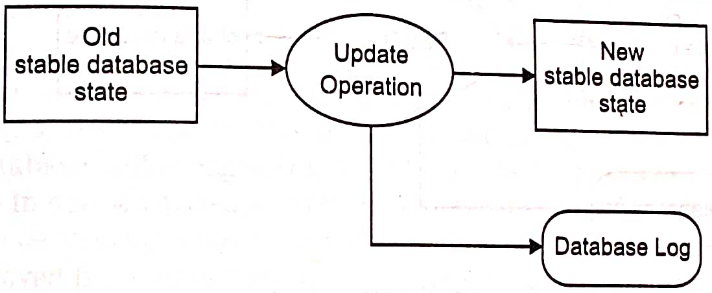


Figure 12.7. Update Operation Execution

Consider the following scenario. The DBMS began executing at time 0 and at time t a system failure occurs. During the period $[0, t]$, two transactions (say, T_1 and T_2) pass through the DBMS, one of which (T_1) has completed (i.e., committed), while the other one has not (see Figure 12.8). The durability property of transactions would require that the effects of T_1 be reflected in the stable database. Similarly, the atomicity property would require that the stable database not contain any of the effects of T_2. However, special precautions need to be taken to ensure this.

Let us assume that the LRM and buffer manager algorithms are such that the buffer pages are written back to the stable database only when the buffer manager needs new buffer space. In other words, the flush command is not used by the LRM and the decision to write back the pages into the stable database is taken at the discretion of the buffer manager. In this case it is possible that the volatile database pages that have been updated by T_1 may not have been written back to the stable database at the time of the failure. Therefore, upon recovery, it is important to be

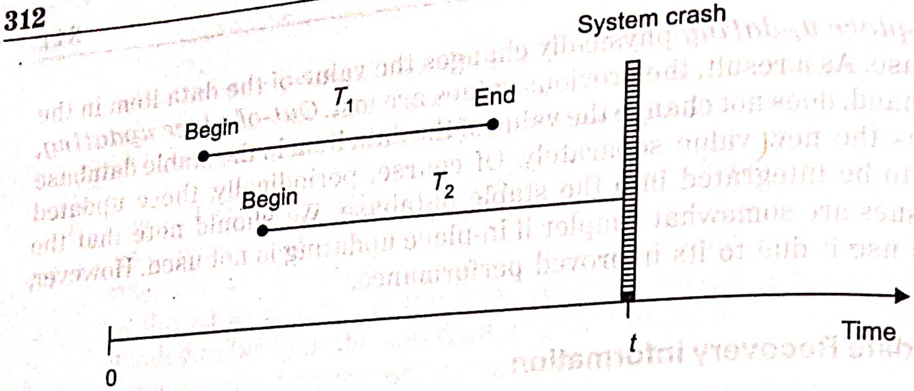


Figure 12.8. Occurrence of a System Failure

able to *redo* the operations of T_1. This requires some information to be stored in the database log about the effects of T_1. Given this information, it is possible to recover the database from its "old" state to the "new" state that reflects the effects of T_2 (Figure 12.9).

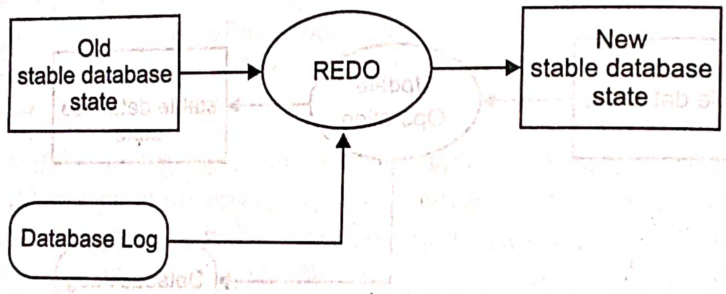


Figure 12.9. REDO Action

Similarly, it is possible that the buffer manager may have had to write into the stable database some of the volatile database pages that have been updated by T_2. Upon recovery from failures it is necessary to *undo* the operations of T_2.[4] Thus the recovery information should include sufficient data to permit the undo by taking the "new" database state that reflects partial effects of T_2, and recovers the "old" state that existed at the start of T_2 (Figure 12.10).

We should indicate that the undo and redo actions are assumed to be idempotent. In other words, their repeated application to a transaction would be equivalent to performing them once. Furthermore, the undo/redo actions form the basis of different methods of executing the commit commands. We discuss this further in Section 12.4.3.

[4]One might think that it could be possible to continue with the operation of T_2 following restart instead of undoing its operations. However, in general it may not be possible for the LRM to determine the point at which the transaction needs to be restarted. Furthermore, the failure may not be a system failure but a transaction failure (i.e., T_2 may actually abort itself) after some of its actions have been reflected in the stable database. Therefore, the possibility of undoing is necessary.

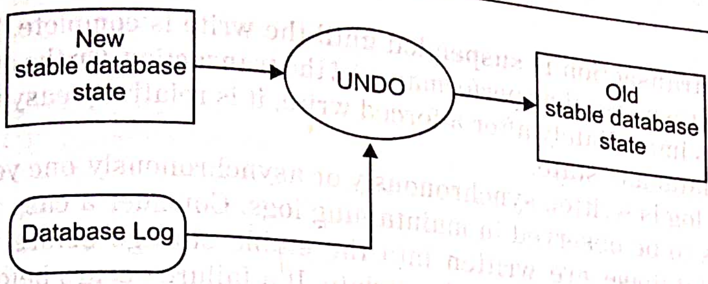


Figure 12.10. UNDO Action

The contents of the log may differ according to the implementation. However, the following minimal information for each transaction is contained in almost all database logs: a begin_transaction record, the value of the data item before the update (called the *before image*), the updated value of the data item (called the *after image*), and a termination record indicating the transaction termination condition (commit, abort). The granularity of the before and after images may be different, as it is possible to log entire pages or some smaller unit. As an alternative to this form of *state logging*, *operational logging*, as in ARIES [Mohan et al., 1993], may be supported where the operations that cause changes to the database are logged rather than the before and after images.

Similar to the volatile database, the log is also maintained in main memory buffers (called *log buffers*) and written back to stable storage (called *stable log*) similar to the database buffer pages (Figure 12.11). The log pages can be written to stable storage in one of two ways. They can be written *synchronously* (more commonly known as *forcing a log*) where the addition of each log record requires that the log be moved from main memory to stable storage. It can also be written *asynchronously*, where the log is moved to stable storage either at periodic intervals or when the buffer fills up. When the log is written synchronously, the

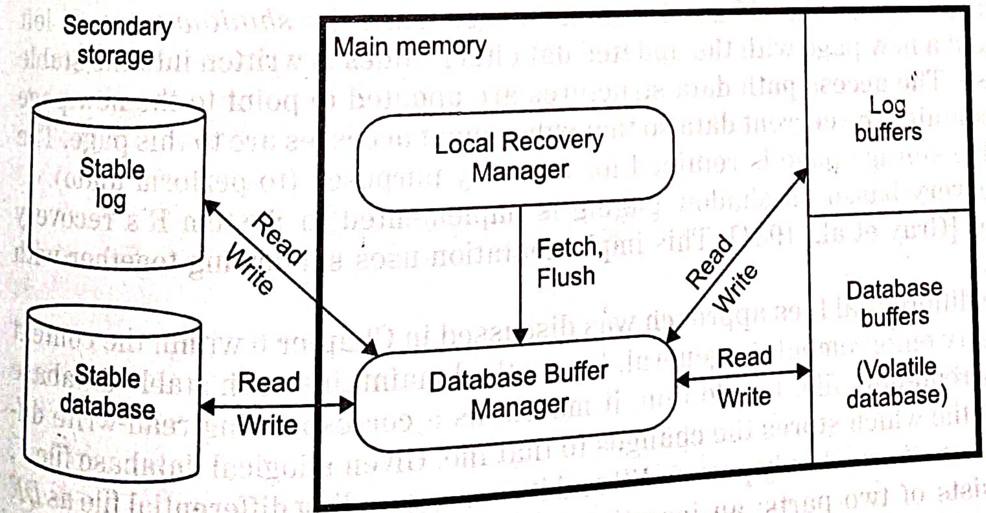


Figure 12.11. Logging Interface

execution of the transaction is suspended until the write is complete. This adds some delay to the response-time performance of the transaction. On the other hand, if a failure occurs immediately after a forced write, it is relatively easy to recover to a consistent database state.

Whether the log is written synchronously or asynchronously, one very important protocol has to be observed in maintaining logs. Consider a case where the updates to the database are written into the stable storage before the log is modified in stable storage to reflect the update. If a failure occurs before the log is written, the database will remain in updated form, but the log will not indicate the update that makes it impossible to recover the database to a consistent and up-to-date state. Therefore, the stable log is always updated prior to the updating of the stable database. This is known as the *write-ahead logging* (*WAL*) protocol [Gray, 1979] and can be precisely specified as follows:

1. Before a stable database is updated (perhaps due to actions of a yet uncommitted transaction), the before images should be stored in the stable log. This facilitates undo.

2. When a transaction commits, the after images have to be stored in the stable log prior to the updating of the stable database. This facilitates redo.

Out-of-Place Update Recovery Information

As we mentioned above, the most common update technique is in-place updating. Therefore, we provide only a brief overview of the other updating techniques and their recovery information. Details can be found in [Verhofstadt, 1978] and the other references given earlier.

Typical techniques for out-of-place updating are *shadowing* ([Astrahan et al., 1979], [Gray, 1979]) and *differential files* [Severence and Lohman, 1976]. Shadowing uses duplicate stable storage pages in executing updates. Thus every time an update is made, the old stable storage page, called the *shadow page*, is left intact and a new page with the updated data item values is written into the stable database. The access path data structures are updated to point to the new page which contains the current data so that subsequent accesses are to this page. The old stable storage page is retained for recovery purposes (to perform undo).

Recovery based on shadow paging is implemented in System R's recovery manager [Gray et al., 1981]. This implementation uses shadowing together with logging.

The differential files approach was discussed in Chapter 6 within the context of integrity enforcement. In general, the method maintains each stable database file as a read-only file. In addition, it maintains a corresponding read-write differential file which stores the changes to that file. Given a logical database file F, let us denote its read-only part as FR and its corresponding differential file as DF. DF consists of two parts: an insertions part, which stores the insertions to F, denoted DF^+, and a corresponding deletions part, denoted DF^-. All updates are treated as the deletion of the old value and the insertion of a new one. Thus each logical file F is considered to be a view defined as $F = (FR \cup DF^{+}) - DF^{-}$. Periodically, the differential file needs to be merged with the read-only base file.

Recovery schemes based on this method simply use private differential files for each transaction, which are then merged with the differential files of each file at commit time. Thus recovery from failures can simply be achieved by discarding the private differential files of noncommitted transactions.

There are studies that indicate that the shadowing and differential files approaches may be advantageous in certain environments. One study by [Agrawal and DeWitt, 1985] investigates the performance of recovery mechanisms based on logging, differential files, and shadow paging, integrated with locking and optimistic (using timestamps) concurrency control algorithms. The results indicate that shadowing, together with locking, can be a feasible alternative to the more common log-based recovery integrated with locking if there are only large (in terms of the base-set size) transactions with sequential access patterns. Similarly, differential files integrated with locking can be a feasible alternative if there are medium-sized and large transactions.

12.4.3 Execution of LRM Commands

Recall that there are five commands that form the interface to the LRM. These are the **begin_transaction**, **read**, **write**, **commit**, and **abort** commands. As we indicated in Chapter 10, some DBMSs do not have an explicit commit command. In this case the end (of transaction) indicator serves as the commit command. For simplicity, we specify commit explicitly.

In this section we introduce a sixth interface command to the LRM: **recover**. The **recover** command is the interface that the operating system has to the LRM. It is used during recovery from system failures when the operating system asks the DBMS to recover the database to the state that existed when the failure occurred.

The execution of some of these commands (specifically, **abort**, **commit**, and **recover**) is quite dependent on the specific LRM algorithms that are used as well as on the interaction of the LRM with the buffer manager. Others (i.e., **begin_transaction**, **read**, and **write**) are quite independent of these considerations.

The fundamental design decision in the implementation of the local recovery manager, the buffer manager, and the interaction between the two components is whether or not the buffer manager obeys the local recovery manager's instructions as to when to write the database buffer pages to stable storage. Specifically, two decisions are involved. The first one is whether the buffer manager may write the buffer pages updated by a transaction into stable storage during the execution of that transaction, or it waits for the LRM to instruct it to write them back. We call this the *fix/no-fix* decision. The reasons for the choice of this terminology will become apparent shortly. Note that it is also called the steal/no-steal decision in [Härder and Reuter, 1983]. The second decision is whether the buffer manager will be forced to flush the buffer pages updated by a transaction into the stable storage at the end of that transaction (i.e., the commit point), or the buffer manager flushes them out whenever it needs to according to its buffer management algorithm. We call this the *flush/no-flush* decision. It is called the force/no-force decision in [Härder and Reuter, 1983].

Accordingly, four alternatives can be identified: (1) no-fix/no-flush, (2) no-fix/flush, (3) fix/no-flush, and (4) fix/flush. We will consider each of these in more detail. However, first we present the execution methods of the **begin_transaction**, **read**, and **write** commands, which are quite independent of these considerations. Where modifications are required in these methods due to different LRM and buffer manager implementation strategies, we will indicate them.

Begin_transaction, Read, and Write Commands

Begin_transaction. This command causes various components of the DBMS to carry out some bookkeeping functions. We will also assume that it causes the LRM to write a begin_transaction record into the log. This is an assumption made for convenience of discussion; in reality, writing of the begin_transaction record may be delayed until the first write to improve performance by reducing I/O.

Read. The **read** command specifies a data item. The LRM tries to read the specified data item from the buffer pages that belong to the transaction. If the data item is not in one of these pages, it issues a **fetch** command to the buffer manager in order to make the data available. Upon reading the data, the LRM returns it to the scheduler.

Write. The **write** command specifies the data item and the new value. As with a read command, if the data item is available in the buffers of the transaction, its value is modified in the database buffers (i.e., the volatile database). If it is not in the private buffer pages, a fetch command is issued to the buffer manager, and the data is made available and updated. The before image of the data page, as well as its after image, are recorded in the log. The local recovery manager then informs the scheduler that the operation has been completed successfully.

No-fix/No-flush

This type of LRM algorithm is called a redo/undo algorithm in [Bernstein et al., 1987] since it requires, as we will see, performing both the redo and undo operations upon recovery. It is called steal/no-force in [Härder and Reuter, 1983].

Abort. As we indicated before, abort is an indication of transaction failure. Since the buffer manager may have written the updated pages into the stable database, abort will have to undo the actions of the transaction. Therefore, the LRM reads the log records for that specific transaction and replaces the values of the updated data items in the volatile database with their before images. The scheduler is then informed about the successful completion of the abort action. This process is called the *transaction undo* or *partial undo*.

An alternative implementation is the use of an *abort list*, which stores the identifiers of all the transactions that have been aborted. If such a list is used, the abort action is considered to be complete as soon as the transaction's identifier is included in the abort list.

Note that even though the values of the updated data items in the stable database are not restored to their before images, the transaction is considered to be aborted at this point. The buffer manager will write the "corrected" volatile database pages into the stable database at a future time, thereby restoring it to its state prior to that transaction.

Commit. The **commit** command causes an end_of_transaction record to be written into the log by the LRM. Under this scenario, no other action is taken in executing a commit command other than informing the scheduler about the successful completion of the commit action.

An alternative to writing an end_of_transaction record into the log is to add the transaction's identifier to a *commit list*, which is a list of the identifiers of transactions that have committed. In this case the commit action is accepted as complete as soon as the transaction identifier is stored in this list.

Recover. The LRM starts the recovery action by going to the beginning of the log and redoing the operations of each transaction for which both a begin_transaction and an end_of_transaction record is found. This is called *partial redo*. Similarly, it undoes the operations of each transaction for which a begin_transaction record is found in the log without a corresponding end_of_transaction record. This action is called *global undo*, as opposed to the transaction undo discussed above. The difference is that the effects of all incomplete transactions need to be rolled back, not one.

If commit list and abort list implementations are used, the recovery action consists of redoing the operations of all the transactions in the commit list and undoing the operations of all the transactions in the abort list. In the remainder of this chapter we will not make this distinction, but rather will refer to both of these recovery implementations as global undo.

No-fix/Flush

The LRM algorithms that use this strategy are called undo/no-redo in [Bernstein et al., 1987] and steal/force in [Härder and Reuter, 1983].

Abort. The execution of **abort** is identical to the previous case. Upon transaction failure, the LRM initiates a partial undo for that particular transaction.

Commit. The LRM issues a **flush** command to the buffer manager, forcing it to write back all the updated volatile database pages into the stable database. The commit command is then executed either by placing a record in the log or by insertion of the transaction identifier into the commit list as specified for the previous case. When all of this is complete, the LRM informs the scheduler that the commit has been carried out successfully.

Recover. Since all the updated pages are written into the stable database at the commit point, there is no need to perform redo; all the effects of successful transactions will have been reflected in the stable database. Therefore, the recovery action initiated by the LRM consists of a global undo.

Fix/No-flush

In this case the LRM controls the writing of the volatile database pages into stable storage. The key here is not to permit the buffer manager to write any updated volatile database page into the stable database until at least the transaction commit point. This is accomplished by the **fix** command, which is a modified version of the **fetch** command whereby the specified page is fixed in the database buffer and cannot be written back to the stable database by the buffer manager. Thus any **fetch** command to the buffer manager for a write operation is replaced by a **fix** command.[5] Note that this precludes the need for a global undo operation and is therefore called a redo/no-undo algorithm in [Bernstein et al., 1987] and a no-force/no-steal algorithm in [Härder and Reuter, 1983].

Abort. Since the volatile database pages have not been written to the stable database, no special action is necessary. To release the buffer pages that have been fixed by the transaction, however, it is necessary for the LRM to send an **unfix** command to the buffer manager for all such pages. It is then sufficient to carry out the abort action either by writing an abort record in the log or by including the transaction in the abort list, informing the scheduler and then forgetting about the transaction.

Commit. The LRM sends an **unfix** command to the buffer manager for every volatile database page that was previously fixed by that transaction. Note that these pages may now be written back to the stable database at the discretion of the buffer manager. The commit command is then executed either by placing an end_of_transaction record in the log or by inserting the transaction identifier into the commit list as specified for the preceding case. When all of this is complete, the LRM informs the scheduler that the commit has been successfully carried out.

Recover. As we mentioned above, since the volatile database pages that have been updated by ongoing transactions are not yet written into the stable database, there is no necessity for global undo. The LRM, therefore, initiates a partial redo action to recover those transactions that may have already committed, but whose volatile database pages may not have yet written into the stable database.

Fix/Flush

This is the case where the LRM forces the buffer manager to write the updated

[5] Of course, any page that was previously fetched for read but is now being updated also needs to be fixed.

volatile database pages into the stable database at precisely the commit point—not before and not after. This strategy is called no-undo/no-redo in [Bernstein et al., 1987] and no-steal/force in [Härder and Reuter, 1983].

Abort. The execution of abort is identical to that of the fix/no-flush case.

Commit. The LRM sends an unfix command to the buffer manager for every volatile database page that was previously fixed by that transaction. It then issues a **flush** command to the buffer manager, forcing it to write back all the unfixed volatile database pages into the stable database.[6] Finally, the **commit** command is processed by either writing an end_of_transaction record into the log or by including the transaction in the commit list. The important point to note here is that all three of these operations have to be executed as an atomic action. One step that can be taken to achieve this atomicity is to issue only a **flush** command, which serves to unfix the pages as well. This eliminates the need to send two messages from the LRM to the buffer manager, but does not eliminate the requirement for the atomic execution of the flush operation and the writing of the database log. The LRM then informs the scheduler that the **commit** has been carried out successfully. Methods for ensuring this atomicity are beyond the scope of our discussion (see [Bernstein et al., 1987]).

Recover. The **recover** command does not need to do anything in this case. This is true since the stable database reflects the effects of all the successful transactions and none of the effects of the uncommitted transactions.

local reliability Protocols (lrm).

12.4.4 Checkpointing

In most of the LRM implementation strategies, the execution of the recovery action requires searching the entire log. This is a significant overhead because the LRM is trying to find all the transactions that need to be undone and redone. The overhead can be reduced if it is possible to build a wall which signifies that the database at that point is up-to-date and consistent. In that case, the redo has to start from that point on and the undo only has to go back to that point. This process of building the wall is called *checkpointing*.

Checkpointing is achieved in three steps [Gray, 1979]:

1. Write a begin_checkpoint record into the log.
2. Collect the checkpoint data into the stable storage.
3. Write an end_checkpoint record into the log.

The first and the third steps enforce the atomicity of the checkpointing operation. If a system failure occurs during checkpointing, the recovery process will not find an end-checkpoint record and will consider checkpointing not completed.

[6]Our discussion here gives the impression that two commands (*unfix* and *flush*) need to be sent to the BM by the LRM for each commit action. We have chosen to explain the action in this way only because of pedagogical simplicity. In reality, it is, of course, preferable to implement one command that instructs the BM to both unfix and flush. Using one command would reduce the message overhead between DBMS components.

There are a number of different alternatives for the data that is collected in Step 2, how it is collected, and where it is stored. We will consider one example here, called *transaction-consistent checkpointing* ([Gray, 1979], [Gray et al., 1981]). The checkpointing starts by writing the begin_checkpoint record in the log and stopping the acceptance of any new transactions by the LRM. Once the active transactions are all completed, all the updated volatile database pages are flushed to the stable database followed by the insertion of an end-checkpoint record into the log. In this case, the redo action only needs to start from the end-checkpoint record in the log. The undo action can go the reverse direction, starting from the end of the log and stopping at the end_checkpoint record.

Transaction-consistent checkpointing is not the most efficient algorithm, since a significant delay is experienced by all the transactions. There are alternative checkpointing schemes such as action-consistent checkpoints, fuzzy checkpoints, and others ([Gray, 1979], [Lindsay, 1979]).

12.4.5 Handling Media Failures

As we mentioned before, the previous discussion on centralized recovery considered nonmedia failures, where the database as well as the log stored in the stable storage survive the failure. Media failures may either be quite catastrophic, causing the loss of the stable database or of the stable log, or they can simply result in partial loss of the database or the log (e.g., loss of a track or two).

The methods that have been devised for dealing with this situation are again based on duplexing. To cope with catastrophic media failures, an *archive* copy of both the database and the log is maintained on a different (tertiary) storage medium, which is typically the magnetic tape or CD-ROM. Thus the DBMS deals with three levels of memory hierarchy: the main memory, random access disk storage, and magnetic tape (Figure 12.12). To deal with less catastrophic failures, having duplicate copies of the database and log may be sufficient.

When a media failure occurs, the database is recovered from the archive copy by redoing and undoing the transactions as stored in the archive log. The real question is how the archive database is stored. If we consider the large sizes of current databases, the overhead of writing the entire database to tertiary storage is significant. Two methods that have been proposed for dealing with this are to perform the archiving activity concurrent with normal processing and to archive the database incrementally as changes occur so that each archive version contains only the changes that have occurred since the previous archiving.

12.5 DISTRIBUTED RELIABILITY PROTOCOLS

As with local reliability protocols, the distributed versions aim to maintain the atomicity and durability of distributed transactions that execute over a number of databases. The protocols address the distributed execution of the **begin_transaction**, **read**, **write**, **abort**, **commit**, and **recover** commands.

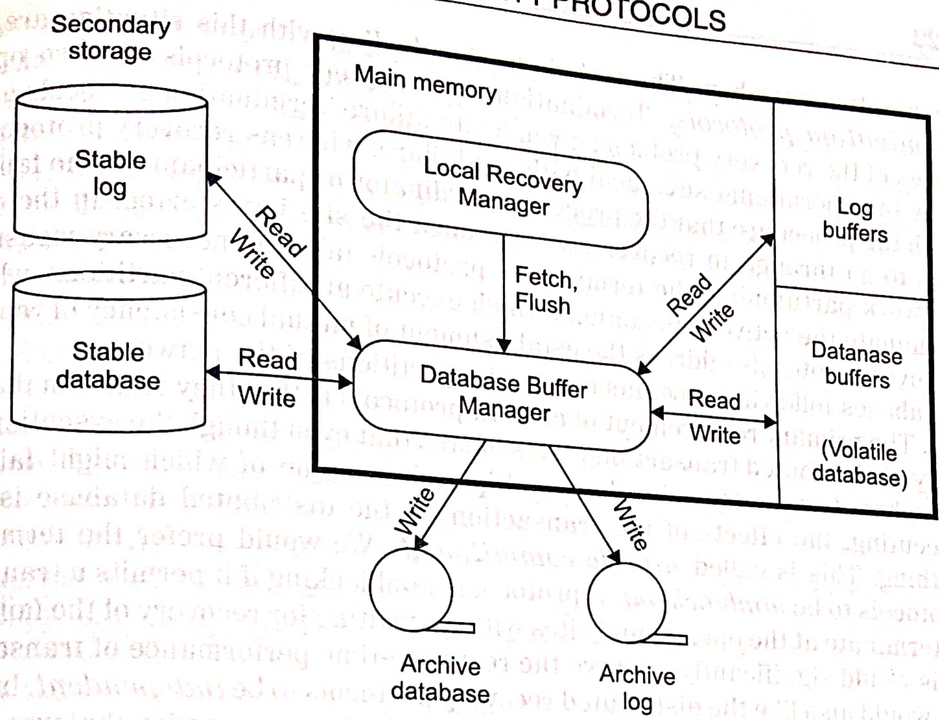


Figure 12.12. Full Memory Hierarchy Managed by LRM and BM

At the outset we should indicate that the execution of the **begin_transaction**, **read**, and **write** commands does not cause any significant problems. **Begin_-transaction** is executed in exactly the same manner as in the centralized case by the transaction manager at the originating site of the transaction. The **read** and **write** commands are executed according to the ROWA rule discussed in Chapter 11. At each of these sites, the commands are executed in the manner described in Section 12.4.3. Similarly, abort is executed by undoing its effects.

To facilitate the description of the distributed reliability protocols, we resort to a commonly used abstraction. We assume that at the originating site of a transaction there is a process that executes its operations. This process is called the *coordinator*. The coordinator communicates with *participant* processes at the other sites which assist in the execution of the transaction's operations. Later we will return to our architectural model and discuss how the coordinator and participant processes can be implemented within that framework.

12.5.1 Components of Distributed Reliability Protocols

The reliability techniques in distributed database systems consist of commit, termination, and recovery protocols. Recall from the preceding section that the commit and recovery protocols specify how the commit and the recover commands are executed. Both of these commands need to be executed differently in a distributed DBMS than in a centralized DBMS. Termination protocols are unique to distributed systems. Assume that during the execution of a distributed transaction, one of the sites involved in the execution fails; we would like the other sites to terminate the transaction somehow. The techniques for dealing with this situation are called *termination protocols*. Termination and recovery protocols are two opposite faces of the recovery problem: given a site failure, termination protocols address how the operational sites deal with the failure, whereas recovery protocols deal with the procedure that the process (coordinator or participant) at the failed site has to go through to recover its state once the site is restarted. In the case of network partitioning, the termination protocols take the necessary measures to terminate the active transactions which execute at different partitions, while the recovery protocols address the establishment of mutual consistency of replicated databases following reconnection of the partitions of the network.

The primary requirement of commit protocols is that they maintain the atomicity of distributed transactions. This means that even though the execution of the distributed transaction involves multiple sites, some of which might fail while executing, the effects of the transaction on the distributed database is all-or-nothing. This is called *atomic commitment*. We would prefer the termination protocols to be *nonblocking*. A protocol is nonblocking if it permits a transaction to terminate at the operational sites without waiting for recovery of the failed site. This would significantly improve the response-time performance of transactions. We would also like the distributed recovery protocols to be *independent*. Independent recovery protocols determine how to terminate a transaction that was executing at the time of a failure without having to consult any other site. Existence of such protocols would reduce the number of messages that need to be exchanged during recovery. Note that the existence of independent recovery protocols would imply the existence of nonblocking termination protocols.[7]

12.5.2 Two-Phase Commit Protocol

Two-phase commit (2PC) is a very simple and elegant protocol that ensures the atomic commitment of distributed transactions. It extends the effects of local atomic commit actions to distributed transactions by insisting that all sites involved in the execution of a distributed transaction agree to commit the transaction before its effects are made permanent. There are a number of reasons why such synchronization among sites is necessary. First, depending on the type of concurrency control algorithm that is used, some schedulers may not be ready to terminate a transaction. For example, if a transaction has read a value of a data item that is updated by another transaction that has not yet committed, the associated scheduler may not want to commit the former. Of course, strict concurrency control algorithms that avoid cascading aborts would not permit the updated value of a data item to be read by any other transaction until the updating transaction terminates. This is sometimes called the *recoverability condition* ([Hadzilacos, 1988], [Bernstein et al., 1987]).

Another possible reason why a participant may not agree to commit is due to deadlocks that require a participant to abort the transaction. Note that in this case the participant should be permitted to abort the transaction without being told to do so. This capability is quite important and is called *unilateral abort*.

[7]The reverse implication is not true, however.

A brief description of the 2PC protocol that does not consider failures is as follows. Initially, the coordinator writes a begin_commit record in its log, sends a "prepare" message to all participant sites, and enters the WAIT state. When a participant receives a "prepare" message, it checks if it could commit the transaction. If so, the participant writes a ready record in the log, sends a "vote-commit" message to the coordinator, and enters READY state; otherwise, the participant writes an abort record and sends a "vote-abort" message to the coordinator. If the decision of the site is to abort, it can forget about that transaction, since an abort decision serves as a veto (i.e., unilateral abort). After the coordinator has received a reply from every participant, it decides whether to commit or to abort the transaction. If even one participant has registered a negative vote, the coordinator has to abort the transaction globally. So it writes an abort record, sends a "global-abort" message to all participant sites, and enters the ABORT state; otherwise, it writes a commit record, sends a "global-commit" message to all participants, and enters the COMMIT state. The participants either commit or abort the transaction according to the coordinator's instructions and send back an acknowledgment, at which point the coordinator terminates the transaction by writing an end_of_transaction record in the log.

Note the manner in which the coordinator reaches a global termination decision regarding a transaction. Two rules govern this decision, which, together, are called the *global commit rule*:

1. If even one participant votes to abort the transaction, the coordinator has to reach a global abort decision.

2. If all the participants vote to commit the transaction, the coordinator has to reach a global commit decision.

The operation of the 2PC protocol between a coordinator and one participant in the absence of failures is depicted in Figure 12.13, where the circles indicate the states and the dashed lines indicate messages between the coordinator and the participants. The labels on the dashed lines specify the nature of the message.

A few important points about the 2PC protocol that can be observed from Figure 12.13 are as follows. First, 2PC permits a participant to unilaterally abort a transaction until it has decided to register an affirmative vote. Second, once a participant votes to commit or abort a transaction, it cannot change its vote. Third, while a participant is in the READY state, it can move either to abort the transaction or to commit it, depending on the nature of the message from the coordinator. Fourth, the global termination decision is taken by the coordinator according to the global commit rule. Finally, note that the coordinator and participant processes enter certain states where they have to wait for messages from one another. To guarantee that they can exit from these states and terminate, timers are used. Each process sets its timer when it enters a state, and if the expected message is not received before the timer runs out, the process times out and invokes its timeout protocol (which will be discussed later).

There are a number of different communication paradigms that, can be employed in implementing a 2PC protocol. The one discussed above and depicted in

Figure 12.13 is called a *centralized 2PC* since the communication is only between the coordinator and the participants; the participants do not communicate among themselves. This communication structure, which is the basis of our subsequent discussions in this chapter, is depicted more clearly in Figure 12.14.

Another alternative is *linear 2PC* (also called *nested 2PC* [Gray, 1979]) where participants can communicate with one another. There is an ordering between the sites in the system for the purposes of communication. Let us assume that the

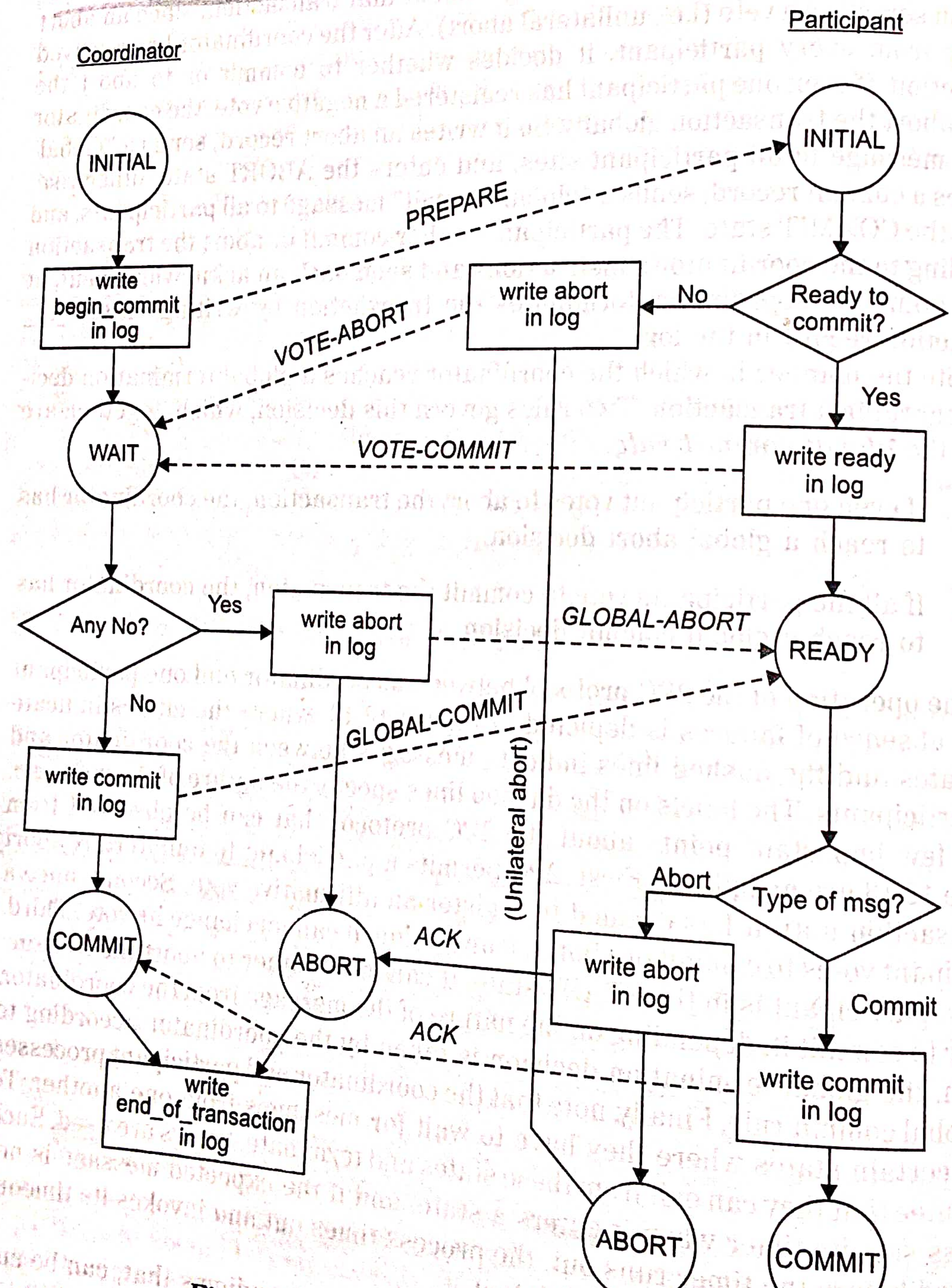


Figure 12.13. 2PC Protocol Actions

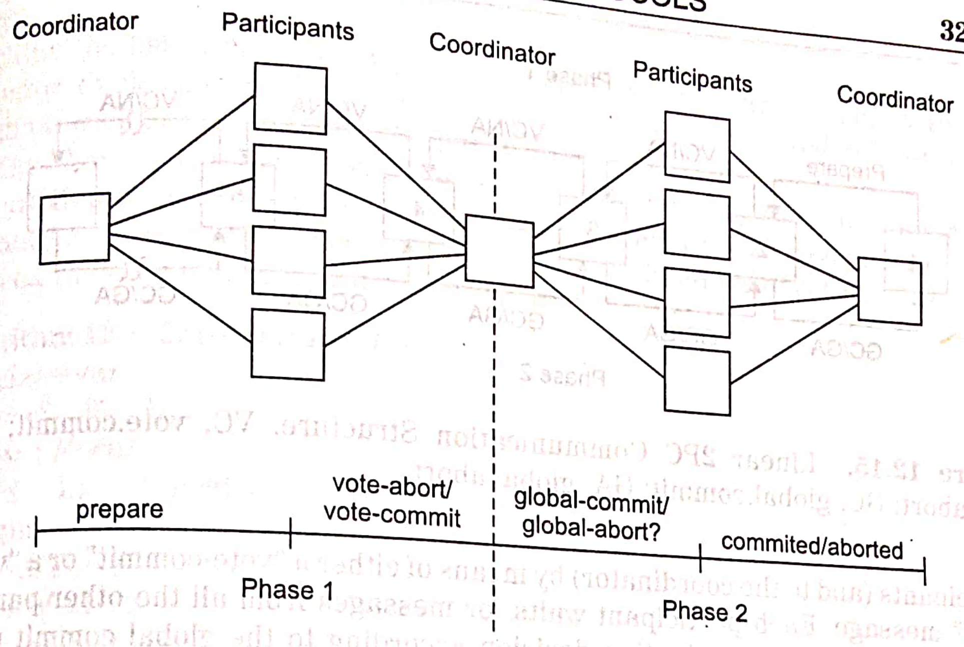


Figure 12.14. Centralized 2PC Communication Structure

ordering among the sites that participate in the execution of a transaction are 1, ..., N, where the coordinator is the first one in the order. The 2PC protocol is implemented by a forward communication from the coordinator (number 1) to N, during which the first phase is completed, and by a backward communication from N to the coordinator, during which the second phase is completed. Thus linear 2PC operates in the following manner.

The coordinator sends the "prepare" message to participant 2. If participant 2 is not ready to commit the transaction, it sends a "vote-abort" message (VA) to participant 3 and the transaction is aborted at this point (unilateral abort by 2). If, on the other hand, participant 2 agrees to commit the transaction, it sends a "vote-commit" message (VC) to participant 3 and enters the READY state. This process continues until a "vote-commit" vote reaches participant N. This is the end of the first phase. If N decides to commit, it sends back *to* $N - 1$ "global-commit" (GC); otherwise, it sends a "global-abort" message (GA). Accordingly, the participants enter the appropriate state (COMMIT or ABORT) and propagate the message back to the coordinator.

Linear 2PC, whose communication structure is depicted in Figure 12.15, incurs fewer messages but does not provide any parallelism. Therefore, it suffers from low response-time performance. It may, however, be suitable for networks that do not have broadcasting capability.

Another popular communication structure for implementation of the 2PC protocol involves communication among all the participants during the first phase of the protocol so that they all independently reach their termination decisions with respect to the specific transaction. This version, called *distributed 2PC*, eliminates the need for the second phase of the protocol since the participants can reach a decision on their own. It operates as follows. The coordinator sends the prepare message to all participants. Each participant then sends its decision to all the other

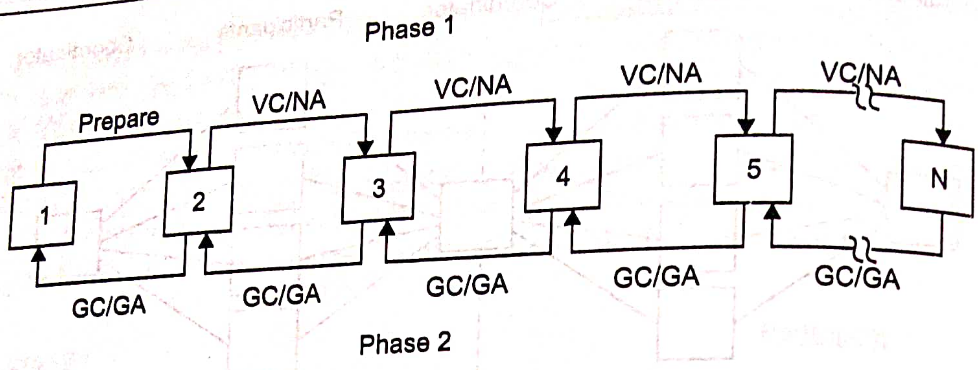


Figure 12.15. Linear 2PC Communication Structure. VC, vote.commit; VA, vote.abort; GC, global.commit; GA, global.abort.

participants (and to the coordinator) by means of either a "vote-commit" or a "vote-abort" message. Each participant waits for messages from all the other participants and makes its termination decision according to the global commit rule. Obviously, there is no need for the second phase of the protocol (someone sending the global abort or global commit decision to the others), since each participant has independently reached that decision at the end of the first phase. The communication structure of distributed commit is depicted in Figure 12.16.

One point that needs to be addressed with respect to the last two versions of 2PC implementation is the following. A participant has to know the identity of either the next participant in the linear ordering (in case of linear 2PC) or of all the participants (in case of distributed 2PC). This problem can be solved by

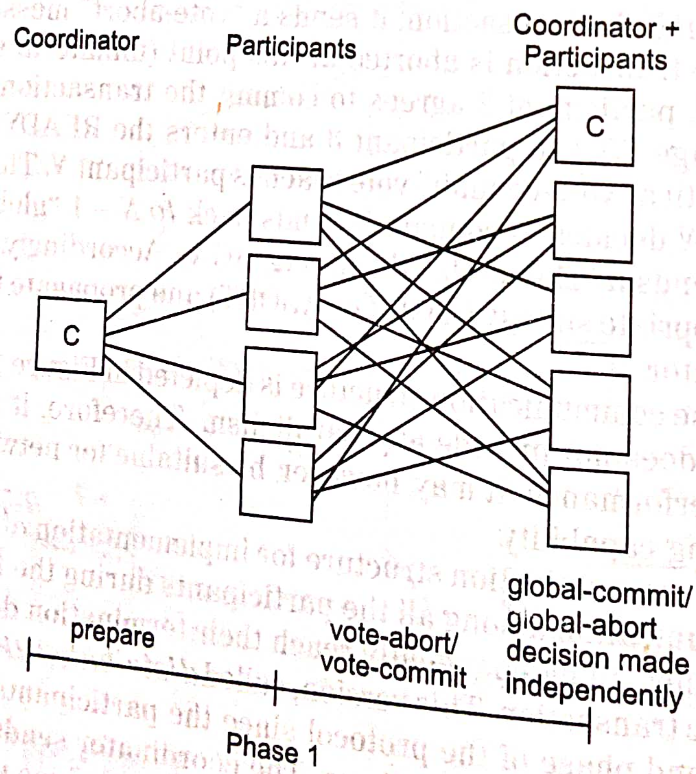


Figure 12.16. Distributed 2PC Communication Structure

attaching the list of participants to the prepare message that is sent by the coordinator. Such an issue does not arise in the case of centralized 2PC since the coordinator clearly knows who the participants are.

The algorithms for the execution of the 2PC protocol by the coordinator and the participants in a centralized communication structure are given in Algorithms 12.1 and 12.2, respectively. The algorithms show the handling of various messages between the coordinator and the participants.

Algorithm 12.1 *2PC-Coordinator*

```
declare-var
    msg : Message
    ev : Event
    PL : List of participant {list compiled prior to the start of 2PC protocol}
begin
    WAIT(ev).
    case of ev                    {possible events are Msg Arrival and Timeout}
      Msg Arrival :
      begin
        Let the arrived message be in msg
        case of msg
          Commit:                            {commit command from scheduler}
          begin
            write begin_commit record in the log
            send "prepare" message to all the participants in PL
            set timer
          end
          Vote-Abort:                        {one participant has voted to abort}
          begin
            write abort record in the log
            send "global-abort" message to all the participants in PL
            set timer
          end
          Vote-Commit:
          begin
            update the list of participants who have answered
            if all the participants have answered then
            begin                            {all must have voted to commit}
              write commit in the log
              send "global-commit" to all the participants in PL
              set timer
            end
          end
          Ack:
          begin
            update the list of participants who have acknowledged
            if all the participants have acknowledged then
```

```
            write end_of_transaction in the log
          else
            send global decision to the unanswering participants
          end-if
        end
      end-case
    end
    Timeout:
    begin
      execute the termination protocol        {this will be discussed later}
    end
  end-case
end. {2PC-Coordinator}
```

Algorithm 12.2 *2PC-Participant*

```
declare-var
  msg : Message
  ev : Event
begin
  WAIT(ev)
  case of ev                    {possible events are MsgArrival and Timeout}
    MsgArrival :
    begin
      Let the arrived message be in msg
      case of msg
        Prepare:
        begin
          if ready to commit then
          begin
            write ready record in the log
            send "vote-commit" message to the coordinator
            set timer
          end
          else begin
            write abort record in the log               {unilateral abort}
            send "vote-abort" message to the oordinator
            call local data processor to abort the transaction
          end
        end-if
        end
        Global-abort:
        begin
          write abort record in the log
          call local data processor to abort the transaction
        end
        Global-commit:
```

```
                begin
                    write commit record in the log
                    call local data processor to commit the transaction
                end
            end-case
        end
        Timeout:
        begin
            execute the termination protocol        {this will be discussed later}
        end
    end-case
end. {2PC-Participant}
```

12.5.3 Variations of 2PC

Two variations of 2PC have been proposed to improve its performance. This is accomplished by reducing (1) the number of messages that are transmitted between the coordinator and the participants, and (2) the number of times logs are written. These protocols are called *presumed abort* and *presumed commit* [Mohan and Lindsay, 1983], [Mohan et al., 1986]. Presumed abort is a protocol that is optimized to handle read-only transactions as well as those update transactions, some of whose processes do not perform any updates to the database while the others do (called partially read-only). The presumed commit protocol is optimized to handle the general update transactions. We will discuss briefly both of these variations.

Presumed Abort 2PC Protocol

In the presumed abort 2PC protocol the following assumption is made. Whenever a prepared participant polls the coordinator about a transaction's outcome and there is no information in virtual storage about it, the response to the inquiry is to abort the transaction. This works since, in the case of a commit, the coordinator does not forget about a transaction until all participants acknowledge, guaranteeing that they will no longer inquire about this transaction.

When this convention is used, it can be seen that the coordinator can forget about a transaction immediately after it decides to abort it. It can write an abort record and not expect the participants to acknowledge the abort command. The coordinator does not need to write an end_of_transaction record after an abort record.

The abort record does not need to be forced, because if a site fails before receiving the decision and then recovers, the recovery routine will check the log to determine the fate of the transaction. Since the abort record is not forced, the recovery routine may not find any information about the transaction, in which case it will ask the coordinator and will be told to abort it. For the same reason, the abort records do not need to be forced by the participants either.

Since it saves some message transmission between the coordinator and the participants in case of aborted transactions, presumed abort 2PC is expected to be more efficient.

Presumed Commit 2PC Protocol

The presumed abort 2PC protocol improves performance by forgetting about transactions once a decision is reached to abort them. Since most transactions are expected to commit, it is reasonable to expect that it may be similarly possible to improve performance for commits. Hence the presumed commit 2PC protocol.

Presumed commit 2PC is based on the premise that if no information about the transaction exists, it should be considered committed. However, it is not an exact dual of presumed abort 2PC, since an exact dual would require that the coordinator forget about a transaction immediately after it decides to commit it, that commit records (also the ready records of the participants) not be forced, and that commit commands need not be acknowledged. Consider, however, the following scenario. The coordinator sends prepared messages and starts collecting information, but fails before being able to collect all of them and reach a decision. In this case, the participants will wait for a while and then turn the transaction over to their recovery routines. Since there is no information about the transaction, the recovery routines of each participant will commit the transaction. The coordinator, on the other hand, will abort the transaction when it recovers, thus causing inconsistency.

A simple variation of this protocol, however, solves the problem and that variant is called the *presumed commit 2PC*. The coordinator, prior to sending the prepare message, force-writes a collecting record, which contains the names of all the participants involved in executing that transaction. The participant then enters the COLLECTING state. It then sends the prepare message and enters the WAIT state. The participants, when they receive the prepare message, decide what they want to do with the transaction, write an abort record, or write a ready record and respond with either a "vote-abort" or a "vote-commit" message. When the coordinator receives decisions from all the participants, it decides to abort or commit the transaction. If the decision is to abort, the coordinator writes an abort record, enters the ABORT state, and sends a "global-abort" message. If it decides to commit the transaction, it writes a commit record, sends a "global-commit" command, and forgets the transaction. When the participants receive a "global-commit" message, they write a commit record and update the database. If they receive a "global-abort" message, they write an abort record and acknowledge. The participant, upon receiving the abort acknowledgment, writes an end record and forgets about the transaction.

12.6 DEALING WITH SITE FAILURES

In this section we consider the failure of sites in the network. Our aim is to develop nonblocking termination and independent recovery protocols. As we indicated before, the existence of independent recovery protocols would imply the existence of nonblocking recovery protocols. However, our discussion addresses both aspects separately. Also note that in the following discussion we consider only the standard 2PC protocol, not its two variants presented above.

Let us first set the boundaries for the existence of nonblocking termination and independent recovery protocols in the presence of site failures. It can formally be proven that such protocols exist when a single site fails. In the case of multiple site failures, however, the prospects are not as promising. An unfortunate result indicates that it is not possible to design independent recovery protocols (and, therefore, nonblocking termination protocols) when multiple sites fail [Skeen and Stonebraker, 1983]. We first develop termination and recovery protocols for the 2PC algorithm and show that 2PC is inherently blocking. We then proceed to the development of atomic commit protocols which are nonblocking in the case of single site failures.

12.6.1 Termination and Recovery Protocols for 2PC

Termination Protocols

The termination protocols serve the timeouts for both the coordinator and the participant processes. A timeout occurs at a destination site when it cannot get an expected message from a source site within the expected time period. In this section we consider that this is due to the failure of the source site.

The method for handling timeouts depends on the timing of failures as well as on the types of failures. We therefore need to consider failures at various points of 2PC execution. This discussion is facilitated by means of the state transition diagram of the 2PC protocol given in Figure 12.17. Note that the state transition diagram is a simplification of Figure 12.13. The states are denoted by circles and the edges represent the state transitions. The terminal states are depicted by concentric circles. The interpretation of the labels on the edges is as follows: the reason for the state transition, which is a received message, is given at the top, and the message that is sent as a result of state transition is given at the bottom.

Coordinator Timeouts. There are three states in which the coordinator can timeout: WAIT, COMMIT, and ABORT. Timeouts during the last two are handled in the same manner. So we need to consider only two cases:

1. *Timeout in the WAIT state.* In the WAIT state, the coordinator is waiting for the local decisions of the participants. The coordinator cannot unilaterally commit the transaction since the global commit rule has not been satisfied. However, it can decide to globally abort the transaction, in which case it writes an abort record in the log and sends a "global-abort" message to all the participants.
2. *Timeout in the COMMIT or ABORT states.* In this case the coordinator is not certain that the commit or abort procedures have been completed

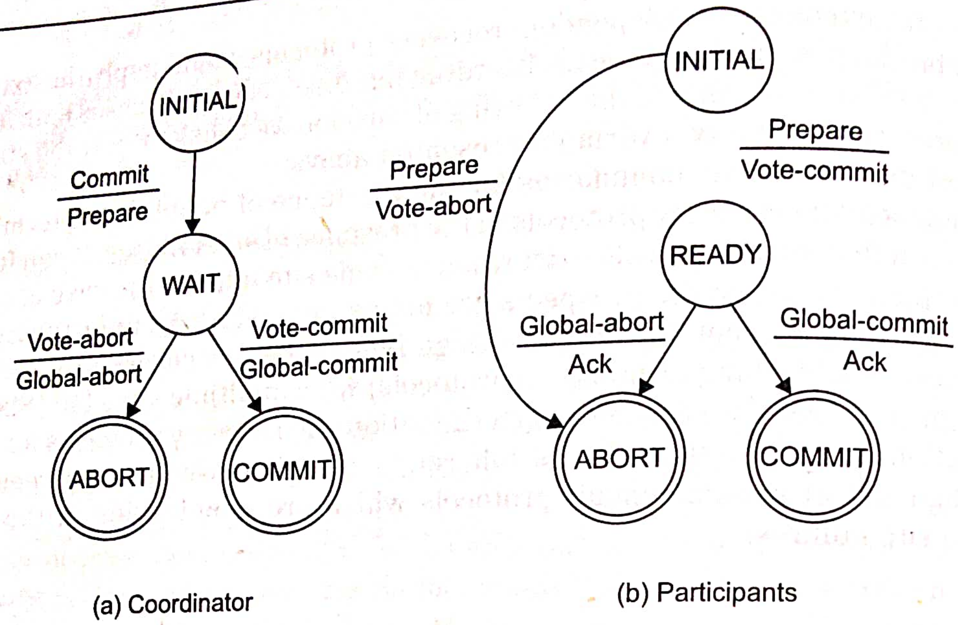


Figure 12.17. State Transitions in 2PC Protocol

by the local recovery managers at all of the participant sites. Thus the coordinator repeatedly sends the "global-commit" or "global-abort" commands to the sites that have not yet responded, and waits for their acknowledgement.

Participant Timeouts. A participant can time out[8] in two states: INITIAL and READY. Let us examine both of these cases.

1. *Timeout in the INITIAL state.* In this state the participant is waiting for a "prepare" message. The coordinator must have failed in the INITIAL state. The participant can unilaterally abort the transaction following a timeout. If the "prepare" message arrives at this participant at a later time, this can be handled in one of two possible ways. Either the participant would check its log, find the abort record, and respond with a "vote-abort," or it can simply ignore the "prepare" message. In the latter case the coordinator would time out in the WAIT state and follow the course we have discussed above.

2. *Timeout in the READY state.* In this state the participant has voted to commit the transaction but does not know the global decision of the coordinator. The participant cannot unilaterally make a decision. Since it is in the READY state, it must have voted to commit the transaction. Therefore, it cannot now change its vote and unilaterally abort it. On the other hand, it cannot unilaterally decide to commit it since it is possible that another participant may have voted to abort it. In this case the participant

[8]In some discussions of the 2PC protocol, it is assumed that the participants do not use timers and do not time out. However, implementing timeout protocols for the participants solves some nasty problems and may speed up the commit process. Therefore, we consider this more general case.

will remain blocked until it can learn from someone (either the coordinator or some other participant) the ultimate fate of the transaction.

Let us consider a centralized communication structure where the participants cannot communicate with one another. In this case the participant that is trying to terminate a transaction has to ask the coordinator for its decision and wait until it receives a response. If the coordinator has failed, the participant will remain blocked. This is undesirable.

If the participants can communicate with each other, a more distributed termination protocol may be developed. The participant that times out can simply ask all the other participants to help it make a decision. Assuming that participant P_i is the one that times out, all the other participants P_j respond in the following manner:

1. P_j is in the INITIAL state. This means that P_j has not yet voted and may not even have received the "prepare" message. It can therefore unilaterally abort the transaction and reply to P_i with a 'Vote-abort" message.

2. P_j is in the READY state. In this state P_j has voted to commit the transaction but has not received any word about the global decision. Therefore, it cannot help P_i to terminate the transaction.

3. P_j is in the ABORT or COMMIT states. In these states, either P_j has unilaterally decided to abort the transaction, or it has received the coordinator's decision regarding global termination. It can, therefore, send P_i either a "vote-commit" or a "vote-abort" message.

Consider how the participant that times out (P_i) can interpret these responses. The following cases are possible:

1. P_i receives "vote-abort" messages from all P_j. This means that none of the other participants had yet voted, but they have chosen to abort the transaction unilaterally. Under these conditions, P_i can proceed to abort the transaction.

2. P_i receives "vote-abort" messages from some P_j, but some other participants indicate that they are in the READY state. In this case P_i can still go ahead and abort the transaction, since according to the global commit rule, the transaction cannot be committed and will eventually be aborted.

3. P_i receives notification from all P_j that they are in the READY state. In this case none of the participants knows enough about the fate of the transaction to terminate it properly.

4. P_i receives "global-abort" or "global-commit" messages from all P_j. In this case all the other participants have received the coordinator's decision. Therefore, P_i can go ahead and terminate the transaction according to the messages it receives from the other participants. Incidentally, note that it is not possible for some of the P_j to respond with a "global-abort" while

others respond with "global-commit" since this cannot be the result of a legitimate execution of the 2PC protocol.

5. P_i receives "global-abort" or "global-commit" from some P_j, whereas others indicate that they are in the READY state. This indicates that some sites have received the coordinator's decision while others are still waiting for it. In this case P_i can proceed as in case 4 above.

These five cases cover all the alternatives that a termination protocol needs to handle. It is not necessary to consider cases where, for example, one participant sends a "vote-abort" message while another one sends "global-commit." This cannot happen in 2PC. During the execution of the 2PC protocol, no process (participant or coordinator) is more than one state transition apart from any other process. For example, if a participant is in the INITIAL state, all other participants are in either the INITIAL or the READY state. Similarly, the coordinator is either in the INITIAL or the WAIT state. Thus all the processes in a 2PC protocol are said to be *synchronous within one state transition* [Skeen, 1981].

Note that in case 3 the participant processes stay blocked, as they cannot terminate a transaction. Under certain circumstances there may be a way to overcome this blocking. If during termination all the participants realize that only the coordinator site has failed, they can elect a new coordinator, which can restart the commit process. There are different ways of electing the coordinator. It is possible either to define a total ordering among all sites and elect the next one in order [Hammer and Shipman, 1980], or to establish a voting procedure among the participants [Garcia-Molina, 1982]. This will not work, however, if both a participant site and the coordinator site fail. In this case it is possible for the participant at the failed site to have received the coordinator's decision and have terminated the transaction accordingly. This decision is unknown to the other participants; thus if they elect a new coordinator and proceed, there is the danger that they may decide to terminate the transaction differently from the participant at the failed site. It is clear that it is not possible to design termination protocols for 2PC that can guarantee nonblocking termination. The 2PC protocol is, therefore, a blocking protocol.

Since we had assumed a centralized communication structure in developing the 2PC algorithms in Algorithms 12.1 and 12.2, we will continue with the same assumption in developing the termination protocols. The portion of code that should be included in the timeout section of the coordinator and the participant 2PC algorithms is given in Algorithms 12.3 and 12.4, respectively.

Algorithm 12.3 *2PC-Coordinator-Terminate*

Timeout:
begin
  if in WAIT state **then**
  begin
    write abort record in the log
    send "global-abort" message to all the participants
  end

```
    else begin
      check for the last log record
      if last log record=abort then
        send "global-abort" to all the participants that have not responded
                                              {coordinator is in ABORT state}
      else
        send "global-commit" to all the participants that have not responded
                                              {coordinator is in COMMIT state}
      end-if
      end
    end-if
    set timer
  end
```

Algorithm 12.4 *2PC-Participant-Terminate*

```
Timeout:
begin
  if in INITIAL state then
    write abort record in the log
  else                                     {participant is in READY state}
    send 'vote-commit" message to the coordinator
    reset timer
  end-if
end
```

Recovery Protocols

In the preceding section we discussed how the 2PC protocol deals with failures from the perspective of the operational sites. In this section we take the opposite viewpoint: we are interested in investigating protocols that a coordinator or participant can use to recover their states when their sites fail and then restart. Remember that we would like these protocols to be independent. However, in general, it is not possible to design protocols that can guarantee independent recovery while maintaining the atomicity of distributed transactions. This is not surprising given the fact that the termination protocols for 2PC are inherently blocking.

In the following discussion we again use the state transition diagram of Figure 12.17. Additionally, we make two interpretive assumptions: (1) the combined action of writing a record in the log and sending a message is assumed to be atomic, and (2) the state transition occurs after the transmission of the response message. For example, if the coordinator is in the WAIT state, this means that it has successfully written the begin_commit record in its log and has successfully transmitted the "prepare" command. This does not say anything, however, about successful completion of the message transmission. Therefore, the "prepare" message may never get to the participants, due to communication failures, which we discuss separately. The first assumption related to atomicity is, of course, unrealistic. However, it simplifies our discussion of fundamental failure cases. At the end of this section we show that the other cases that arise from the relaxation of this assumption can be handled by a combination of the fundamental failure cases.

Coordinator Site Failures. The following cases are possible:

1. *The coordinator fails while in the INITIAL state.* This is before the coordinator has initiated the commit procedure. Therefore, it will start the commit process upon recovery.

2. *The coordinator fails while in the WAIT state.* In this case the coordinator has sent the "prepare" command. Upon recovery, the coordinator will restart the commit process for this transaction from the beginning by sending the "prepare" message one more time.

3. *The coordinator fails while in the COMMIT or ABORT states.* In this case the coordinator will have informed the participants of its decision and terminated the transaction. Thus, upon recovery, it does not need to do anything if all the acknowledgments have been received. Otherwise, the termination protocol is involved.

Participant Site Failures. There are three alternatives to consider:

1. *A participant fails in the INITIAL state.* Upon recovery, the participant should abort the transaction unilaterally. Let us see why this is acceptable. Note that the coordinator will be in the INITIAL or WAIT state with respect to this transaction. If it is in the INITIAL state, it will send a "prepare" message and then move to the WAIT state. Because of the participant site's failure, it will not receive the participant's decision and will time out in that state. We have already discussed how the coordinator would handle timeouts in the WAIT state by globally aborting the transaction.

2. *A participant fails while in the READY state.* In this case the coordinator has been informed of the failed site's affirmative decision about the transaction before the failure. Upon recovery, the participant at the failed site can treat this as a timeout in the READY state and hand the incomplete transaction over to its termination protocol.

3. *A participant fails while in the ABORT or COMMIT state.* These states represent the termination conditions, so, upon recovery, the participant does not need to take any special action.

Additional Cases. Let us now consider the cases that may arise when we relax the assumption related to the atomicity of the logging and message sending actions. In particular, we assume that a site failure may occur after the coordinator or a participant has written a log record but before it can send a message. For this discussion the reader may wish to refer to Figure 12.13.

1. *The coordinator fails after the begin_commit record is written in the log but before the "prepare" command is sent.* The coordinator would react to this as a failure in the WAIT state (case 2 of the coordinator failures discussed above) and send the "prepare" command upon recovery.

2. *A participant site fails after writing the ready record in the log but before sending the "vote-commit" message.* The failed participant sees this as case 2 of the participant failures discussed before.

3. *A participant site fails after writing the abort record in the log but before sending the "vote-abort" message.* This is the only situation that is not covered by the fundamental cases discussed before. However, the participant does not need to do anything upon recovery in this case. The coordinator is in the WAIT state and will time out. The coordinator termination protocol for this state globally aborts the transaction.

4. *The coordinator fails after logging its final decision record (abort or commit), but before sending its "global-abort" or "global-commit" message to the participants.* The coordinator treats this as its case 3, while the participants treat it as a timeout in the READY state.

5. *A participant fails after it logs an abort or a commit record but before it sends the acknowledgment message to the coordinator.* The participant can treat this as its case 3. The coordinator will handle this by timeout in the COMMIT or ABORT state.

12.6.2 Three-Phase Commit Protocol

The three-phase commit protocol (3PC) [Skeen, 1981] is designed as a nonblocking protocol. We will see in this section that it is indeed nonblocking when failures are restricted to site failures.

Let us first consider the necessary and sufficient conditions for designing nonblocking atomic commitment protocols. A commit protocol that is synchronous within one state transition is nonblocking if and only if its state transition diagram contains neither of the following:

1. No state that is "adjacent" to both a commit and an abort state

2. No noncommittable state that is "adjacent" to a commit state ([Skeen, 1981], [Skeen and Stonebraker, 1983])

The term *adjacent* here means that it is possible to go from one state to the other with a single state transition.

Consider the COMMIT state in the 2PC protocol (see Figure 12.17). If any process is in this state, we know that all the sites have voted to commit the transaction. Such states are called *committable*. There are other states in the 2PC protocol that are *noncommittable*. The one we are interested in is the READY state, which is noncommittable since the existence of a process in this state does not imply that all the processes have voted to commit the transaction.

It is obvious that the WAIT state in the coordinator and the READY state in the participant 2PC protocol violate the nonblocking conditions we have stated above. Therefore, one might be able to make the following modification to the 2PC protocol to satisfy the conditions and turn it into a nonblocking protocol.

We can add another state between the WAIT (and READY) and COMMIT states which serves as a buffer state where the process is ready to commit (if that is the final decision) but has not yet committed. The state transition diagrams for the coordinator and the participant in this protocol are depicted in Figure 12.18. This is called the three-phase commit protocol (3PC) because there are three state transitions from the INITIAL state to a COMMIT state. The execution of the protocol between the coordinator and one participant is depicted in Figure 12.19. Note that this is identical to Figure 12.13 except for the addition of the PRECOMMIT state. Observe that 3PC is also a protocol where all the states are synchronous within one state transition. Therefore, the foregoing conditions for nonblocking 2PC apply to 3PC.

It is possible to design different 3PC algorithms depending on the communication topology. The one given in Figure 12.19 is centralized. It is also straightforward to design a distributed 3PC protocol. A linear 3PC protocol is somewhat more involved, so we leave it as an exercise.

Termination Protocol

As we did in discussing the termination protocols for handling timeouts in the 2PC protocol, let us investigate timeouts at each state of the 3PC protocol.

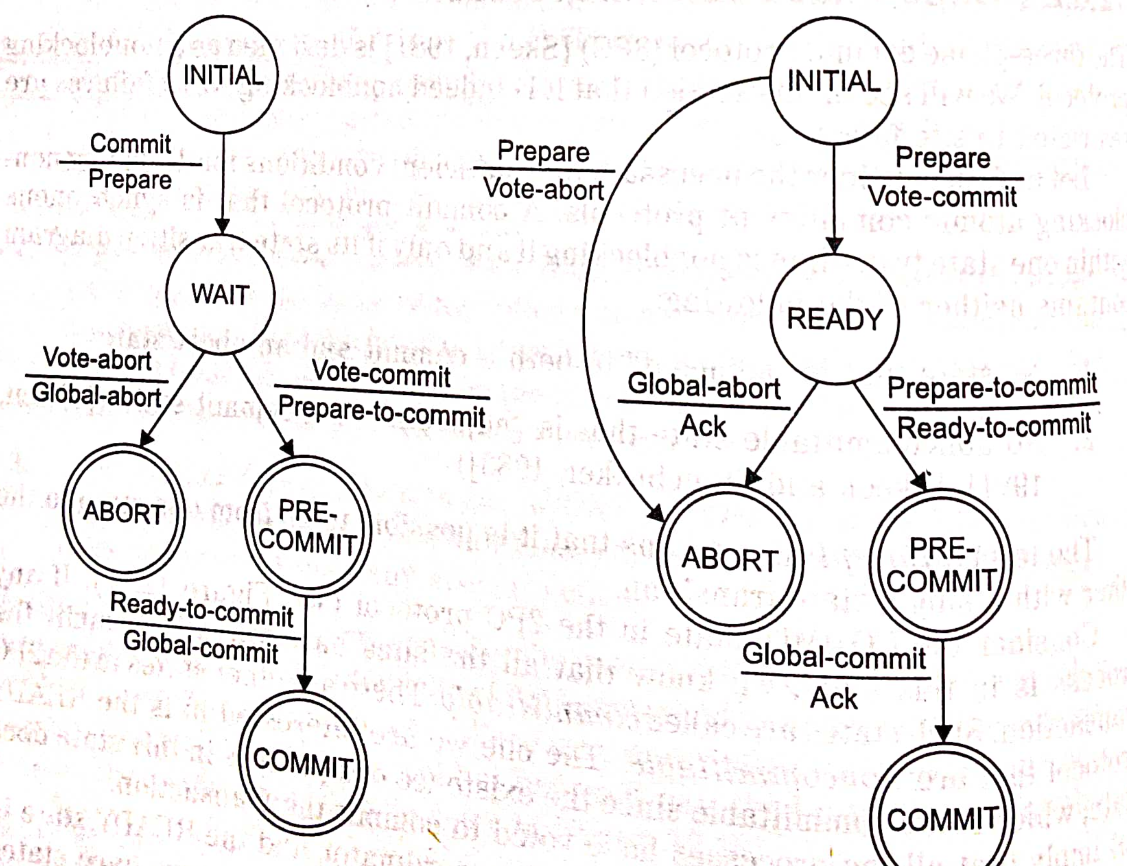


Figure 12.18. State Transitions in 3PC Protocol

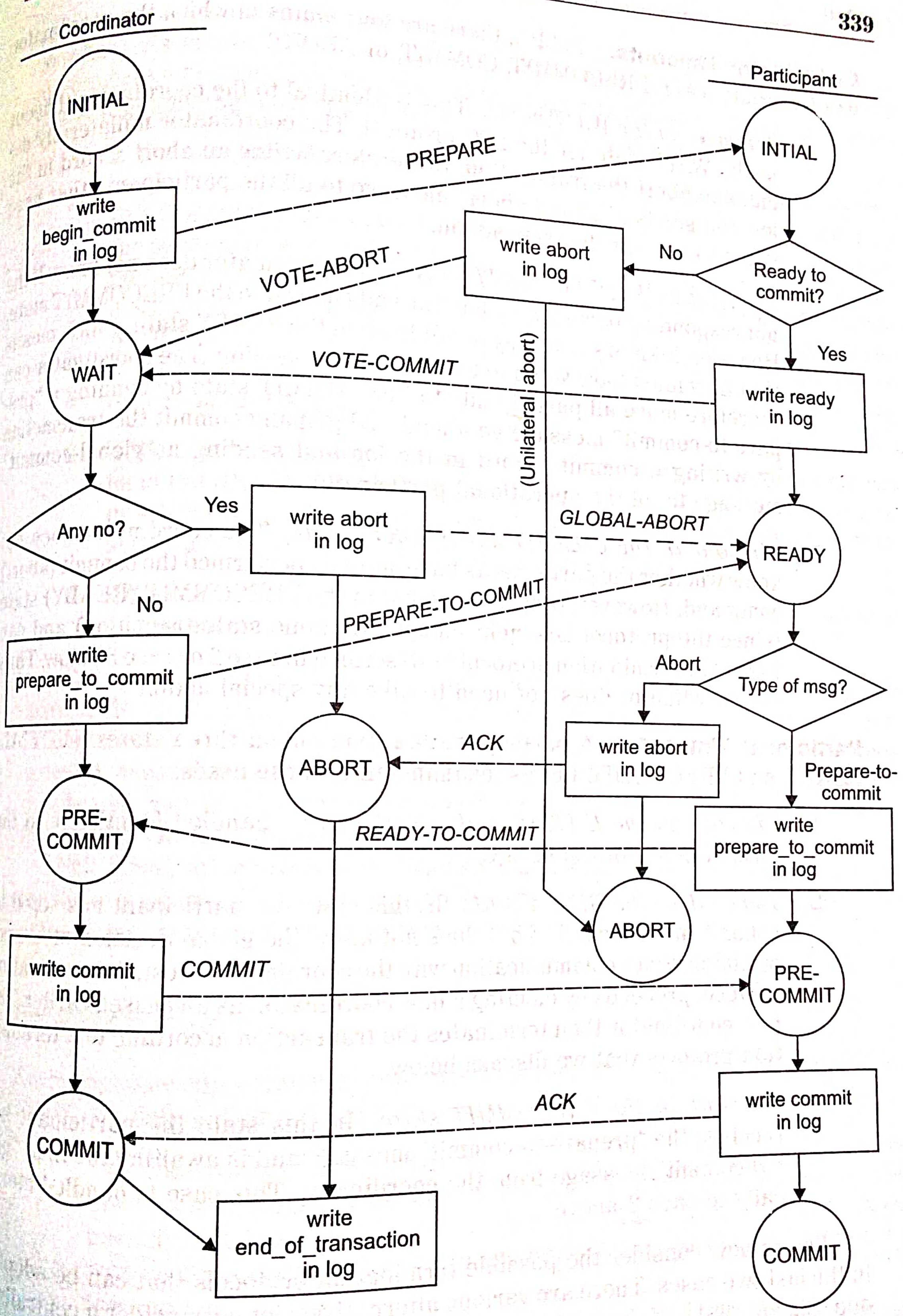


Figure 12.19. 3PC Protocol Actions

Coordinator Timeouts. In 3PC, there are four states in which the coordinator can time out: WAIT, PRECOMMIT, COMMIT, or ABORT.

1. *Timeout in the WAIT state.* This is identical to the coordinator timeout in the WAIT state for the 2PC protocol. The coordinator unilaterally decides to abort the transaction. It therefore writes an abort record in the log and sends a "global-abort" message to all the participants that have voted to commit the transaction.

2. *Timeout in the PRECOMMIT state.* The coordinator does not know if the nonresponding participants have already moved to the PRECOMMIT state. However, it knows that they are at least in the READY state, which means that they must have voted to commit the transaction. The coordinator can therefore move all participants to PRECOMMIT state by sending a "prepare-to-commit" message go ahead and globally commit the transaction by writing a commit record in the log and sending a "global-commit" message to all the operational participants.

3. *Timeout in the COMMIT (or ABORT) state.* The coordinator does not know whether the participants have actually performed the commit (abort) command. However, they are at least in the PRECOMMIT (READY) state (since the protocol is synchronous within one state transition) and can follow the termination protocol as described in case 2 or case 3 below. Thus the coordinator does not need to take any special action.

Participant Timeouts. A participant can time out in three states: INITIAL, READY, and PRECOMMIT. Let us examine all of these cases.

1. *Timeout in the INITIAL state.* This can be handled identically to the termination protocol of 2PC.

2. *Timeout in the READY state.* In this state the participant has voted to commit the transaction but does not know the global decision of the coordinator. Since communication with the coordinator is lost, the termination protocol proceeds by electing a new coordinator, as discussed earlier. The new coordinator then terminates the transaction according to a termination protocol that we discuss below.

3. *Timeout in the PRECOMMIT state.* In this state the participant has received the "prepare-to-commit" message and is awaiting the final "global-commit" message from the coordinator. This case is handled identically to case 2 above.

Let us now consider the possible termination protocols that can be adopted in the last two cases. There are various alternatives; let us consider a centralized one [Skeen, 1981]. We know that the new coordinator can be in one of three states: WAIT, PRECOMMIT, COMMIT or ABORT. It sends its own state to all the operational participants, asking them to assume that state. Any participant who has proceeded ahead of the new coordinator (which is possible since it may have already received

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1. If the new coordinator is in the WAIT state, it will globally abort the transaction. The participants can be in the INITIAL, READY, ABORT, or PRECOMMIT states. In the first three cases, there is no problem. However, the participants in the PRECOMMIT state are expecting a "global-commit" message, but they get a "global-abort" instead. Their state transition diagram does not indicate any transition from the PRECOMMIT to the ABORT state. This transition is necessary for the termination protocol, so it should be added to the set of legal transitions that can occur during execution of the termination protocol.

2. If the new coordinator is in the PRECOMMIT state, the participants can be in the READY, PRECOMMIT or COMMIT states. No participant can be in ABORT state. The coordinator will therefore globally commit the transaction and send a "global-commit" message.

3. If the new coordinator is in the ABORT state, at the end of the first message all the participants will have moved into the ABORT state as well.

The new coordinator is not keeping track of participant failures during this process. It simply guides the operational sites toward termination. If some participants fail in the meantime, they will have to terminate the transaction upon recovery according to the methods discussed in the next section. Also, the new coordinator may fail during the process; the termination protocol therefore needs to be reentrant in implementation.

This termination protocol is obviously nonblocking. The operational sites can properly terminate all the ongoing transactions and continue their operations. The proof of correctness of the algorithm is given in [Skeen, 1982a].

Recovery Protocols

There are some minor differences between the recovery protocols of 3PC and those of 2PC. We only indicate those differences.

1. *The coordinator fails while in the WAIT state.* This is the case we discussed at length in the earlier section on termination protocols. The participants have already terminated the transaction. Therefore, upon recovery, the coordinator has to ask around to determine the fate of the transaction.

2. *The coordinator fails while in the PRECOMMIT state.* Again, the termination protocol has guided the operational participants toward termination. Since it is now possible to move from the PRECOMMIT state to the

ABORT state during this process, the coordinator has to ask around to determine the fate of the transaction.

3. *A participant fails while in the PRECOMMIT state.* It has to ask around to determine how the other participants have terminated the transaction.

One property of the 3PC protocol becomes obvious from this discussion. When using the 3PC protocol, we are able to terminate transactions without blocking. However, we pay the price that fewer cases of independent recovery are possible. This also results in more messages being exchanged during recovery.

12.7 NETWORK PARTITIONING

In this section we consider how the network partitions can be handled by the atomic commit protocols that we discussed in the preceding section. Network partitions are due to communication line failures and may cause the loss of messages, depending on the implementation of the communication subnet. A partitioning is called a *simple partitioning* if the network is divided into only two components; otherwise, it is called *multiple partitioning*.

The termination protocols for network partitioning address the termination of the transactions that were active in each partition at the time of partitioning. If one can develop nonblocking protocols to terminate these transactions, it is possible for the sites in each partition to reach a termination decision (for a given transaction) which is consistent with the sites in the other partitions. This would imply that the sites in each partition can continue executing transactions despite the partitioning.

Unfortunately, it is not in general possible to find nonblocking termination protocols in the presence of network partitions. Remember that our expectations regarding the reliability of the communication subnet are minimal. If a message cannot be delivered, it is simply lost. In this case it can be proven that no nonblocking atomic commitment protocol exists that is resilient to network partitioning [Skeen and Stonebraker, 1983]. This is quite a negative result since it also means that if network partitioning occurs, we cannot continue normal operations in all partitions, which limits the availability of the entire distributed database system. A positive counter result, however, indicates that it is possible to design nonblocking atomic commit protocols that are resilient to simple partitions. Unfortunately, if multiple partitions occur, it is again not possible to design such protocols [Skeen and Stonebraker, 1983].

In the remainder of this section we discuss a number of protocols that address network partitioning in nonreplicated databases. We did not need to make this distinction in our discussion of site failures in the preceding section, since the termination protocols are identical for both types of database organization. However, the same is not true in the case of network partitioning. Recall from Chapter 11 that in the case of replicated databases, the replica control protocol has to be involved in mapping a read or a write on a logical data item to a read or a write on the physical data item copies. In the presence of network partitioning, the copies with the management of network partitioning. In the presence of network partitioning of nonreplicated databases, on the other hand, the major concern is with the termination of transactions that were active at the time of partitioning. Any new transaction that accesses a data item that is stored in another partition is simply blocked and has to await the repair of the network. Concurrent accesses to the data items within one partition can be handled by the concurrency control algorithm. The significant problem, therefore, is to ensure that the transaction terminates properly. In short, in nonreplicated databases, the network partitioning problem is handled by the commit protocol, and more specifically, by the termination and recovery protocols, whereas in replicated databases it is the responsibility of the replica control protocol. Since replica control protocols have more general objectives than simply managing network partitioning, we defer them to the next section.

The absence of nonblocking protocols that would guarantee atomic commitment of distributed transactions points to an important design decision. We can either permit all the partitions to continue their normal operations and accept the fact that database consistency may be compromised, or we guarantee the consistency of the database by employing strategies that would permit operation in one of the partitions while the sites in the others remain blocked. This decision problem is the premise of a classification of partition handling strategies. We can classify the strategies as *pessimistic* or *optimistic* [Davidson et al., 1985]. Pessimistic strategies emphasize the consistency of the database, and would therefore not permit transactions to execute in a partition if there is no guarantee that the consistency of the database can be maintained. Optimistic approaches, on the other hand, emphasize the availability of the database even if this would cause inconsistencies.

The second dimension is related to the correctness criterion. If serializability is used as the fundamental correctness criterion, such strategies are called *syntactic* since the serializability theory uses only syntactic information. However, if we use a more abstract correctness criterion that is dependent on the semantics of the transactions or the database, the strategies are said to be *semantic*.

Consistent with the correctness criterion that we have adopted in this book (serializability), we consider only syntactic approaches in this section. The following two sections outline various syntactic strategies for nonreplicated databases.

All the known termination protocols that deal with network partitioning in the case of nonreplicated databases are pessimistic. Since the pessimistic approaches emphasize the maintenance of database consistency, the fundamental issue that we need to address is which of the partitions can continue normal operations. We consider two approaches.

12.7.1 Centralized Protocols

Centralized termination protocols are based on the centralized concurrency control algorithms discussed in Chapter 11. Recall that these may be of two types: primary site and primary copy. In the case of primary site concurrency control algorithms, it makes sense to permit the operation of the partition that contains the primary site, since it manages the lock tables.

In the case of primary copy concurrency control algorithms, more than one partition may be operational for different queries. For any given query, only the partition that contains the primary copies of the data items that are in the write set of that transaction can execute that transaction.

Both of these are simple approaches that would work well, but they are dependent on the concurrency control mechanism employed by the distributed database manager. Furthermore, they expect each site to be able to differentiate network partitioning from site failures properly. This is necessary since the participants in the execution of the commit protocol react differently to the different types of failures.

12.7.2 Voting-based Protocols

Voting as a technique for managing concurrent data accesses has been proposed by a number of researchers. A straightforward voting with majority was first proposed in [Thomas, 1979] as a concurrency control method for fully replicated databases. The fundamental idea is that a transaction is executed if a majority of the sites vote to execute it.

The idea of majority voting has been generalized to voting with *quorums*. Quorum-based voting can be used as a replica control method (as we discuss in the next section), as well as a commit method to ensure transaction atomicity in the presence of network partitioning. In the case of nonreplicated databases, this involves the integration of the voting principle with commit protocols. We present next a specific proposal along this line [Skeen, 1982a].

Every site in the system is assigned a vote V_i. Let us assume that the total number of votes in the system is V, and the abort and commit quorums are V_a and V_c, respectively. Then the following rules must be obeyed in the implementation of the commit protocol:

1. $V_a + V_c > V$, where $0 \le V_a, V_c \le V$.
2. Before a transaction commits, it must obtain a commit quorum V_c.
3. Before a transaction aborts, it must obtain an abort quorum V_a.

The first rule ensures that a transaction cannot be committed and aborted at the same time. The next two rules indicate the votes that a transaction has to obtain before it can terminate one way or the other.

The integration of these rules into the 3PC protocol requires a minor modification of the third phase. For the coordinator to move from the PRECOMMIT state to the COMMIT state, and to send the "*global-commit*" command, it is necessary for it to have obtained a commit quorum from the participants. This would satisfy rule 2. Note that we do not need to implement rule 3 explicitly. This is due to the fact that a transaction which is in the WAIT or READY state is willing to abort the transaction. Therefore, an abort quorum already exists.

Let us now consider the termination of transactions in the presence of failures. When a network partitioning occurs, the sites in each partition elect a new coordinator, similarly to the 3PC termination protocol in the case of site failures. There is a fundamental difference, however. It is not possible to make the transition from the WAIT or READY state to the ABORT state in one state transition, for a number of reasons. First, more than one coordinator is trying to terminate the transaction. We do not want them to terminate differently or the transaction execution will not be atomic. Therefore, we want the coordinators to obtain an abort quorum explicitly. Second, if the newly elected coordinator fails, it is not known whether a commit or abort quorum was reached. Thus it is necessary that participants make an explicit decision to join either the commit or the abort quorum and not change their votes afterward. Unfortunately, the READY (or WAIT) state does not satisfy these requirements. Thus we introduce another state, PREABORT, between the READY and ABORT states. The transition from the PREABORT state to the ABORT state requires an abort quorum. The state transition diagram is given in Figure 12.20.

With this modification, the termination protocol works as follows. Once a new coordinator is elected, it requests all participants to report their local states. Depending on the responses, it terminates the transaction as follows:

1. If at least one participant is in the COMMIT state, the coordinator decides to commit the transaction and sends a "global-commit" message to all the participants.

2. If at least one participant is in the ABORT state, the coordinator decides to abort the transaction and sends a "global-abort" message to all the participants.

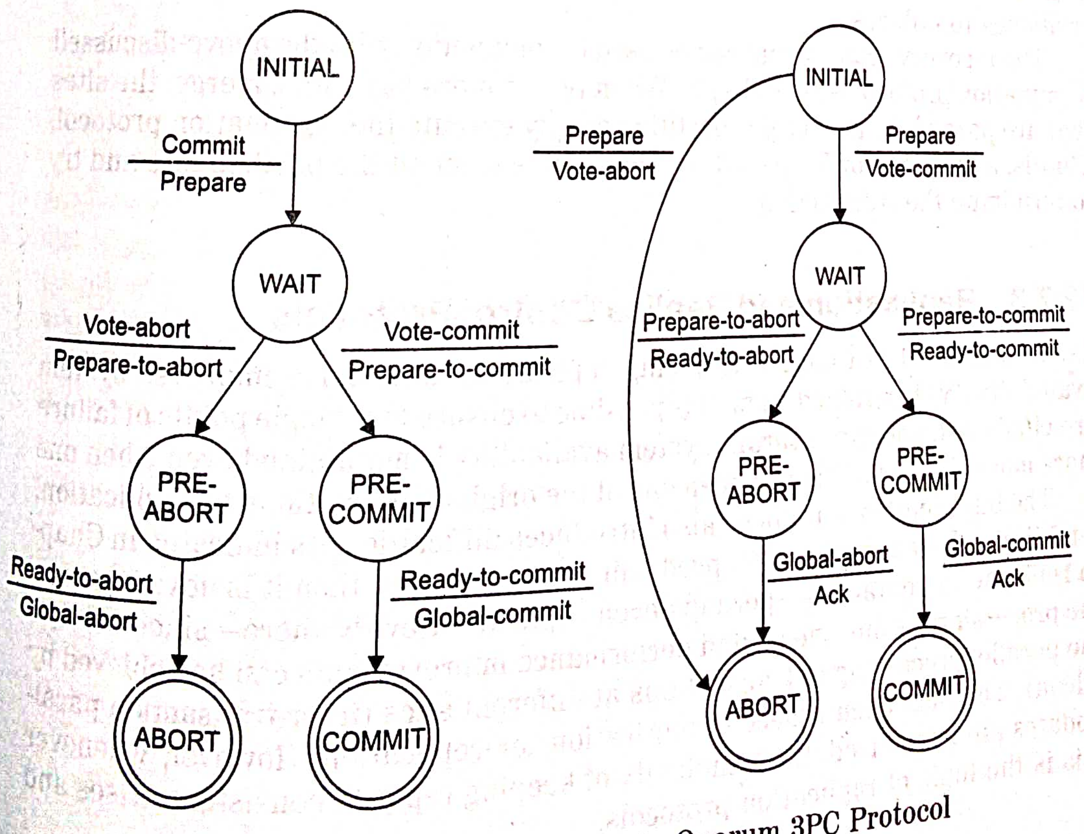

Figure 12.20. State Transitions in Quorum 3PC Protocol

3. If a commit quorum is reached by the votes of participants in the PRECOMMIT state, the coordinator decides to commit the transaction and sends a "global-commit" message to all the participants.

4. If an abort quorum is reached by the votes of participants in the PREABORT state, the coordinator decides to abort the transaction and sends a "global-abort" message to all the participants.

5. If case 3 does not hold but the sum of the votes of the participants in the PRECOMMIT and READY states are enough to form a commit quorum, the coordinator moves the participants to the PRECOMMIT state by sending a "prepare-to-commit" message. The coordinator then waits for case 3 to hold.

6. Similarly, if case 4 does not hold but the sum of the votes of the participants in the PREABORT and READY states are enough to form an abort quorum, the coordinator moves the participants to the PREABORT state by sending a "prepare-to-abort" message. The coordinator then waits for case 4 to hold.

Two points are important about this quorum-based commit algorithm. First, it is blocking; the coordinator in a partition may not be able to form either an abort or a commit quorum if messages get lost or multiple partitionings occur. This is hardly surprising given the theoretical bounds that we discussed previously. The second point is that the algorithm is general enough to handle site failures as well as network partitioning. Therefore, this modified version of 3PC can provide more resiliency to failures.

The recovery protocol that can be used in conjunction with the above-discussed termination protocol is very simple. When two or more partitions merge, the sites that are part of the new larger partition simply execute the termination protocol. That is, a coordinator is elected to collect votes from all the participants and try to terminate the transaction.

12.7.3 Replication and Replica Control Protocols

As we discussed in Chapter 1, having replicas of data items improves system availability. With careful design, it is possible to ensure that single points of failure are eliminated and the "overall" system availability is maintained even when one more more sites fail. This has been one of the original motivations for replication.

The introduction of replicas also introduces difficulties. As indicated in Chapter 5, if the system workload is predominantly read-only, then it is advantageous to replicate the database—perhaps even replicate it everywhere—since most of the processing becomes local and performance improvements can be achieved by the parallel processing of transactions at different sites (inter-transaction parallelism). This has been a second motivation for replication. However, whenever updates are introduced, the complexity of keeping replicas consistent arises and this is the topic of replication protocols.

Even though the original motivations for replication have been to improve system availability and to improve system performance when reads are predominant, it has become fashionable in recent years to propose replication as an alternative to commit protocols. The argument that is made is that 2PC has a high overhead and, therefore, may not be desirable. Instead, data can be replicated at multiple sites and updated independently even if this leads to inconsistent replicas; inconsistencies can be eliminated later by *lazy replication* protocols which update replicas whenever an access occurs. We discuss these protocols later in this section. It is important to note, however, that replication is not a direct alternative to commit protocols; the two have quite different semantics since commit protocols force the database replicas to have the same value (one-copy equivalence) at the end of the update transaction while lazy replication only promises that they will converge to the same value at some later time.

12.7.4 Strict Replica Control Protocols

What we call *strict replica control protocols* are those that enforce one-copy equivalence as the correctness criterion (Chapter 10). To recall, this correctness criterion requires that all the database copies be mutually consistent at the end of each up-date transaction. We have already discussed one protocol for enforcing one-copy equivalence: the ROWA protocol which converts a logical read to a read on any one of the replicas, and converts a logical write to a write an all the replicas. Thus, when the update transaction commits, all of the replicas have the same value.

ROWA is simple and elegant. However, as we saw during the discussion of commit protocols, it has one significant drawback. Even if one of the replicas is unavailable, then the update transaction cannot be terminated. So, ROWA fails in meeting one of the fundamental goals of replication, namely providing higher availability.

An alternative to ROWA which attempts to address the low availability problem is the Read One Write All Available (ROWA-A) protocol. The general idea is that the write commands are executed on all the available copies and the transaction terminates. The copies that were unavailable at the time will have to "catch up" when they become available.

There have been various versions of this protocol. One that we will discuss here is due to [Bernstein and Goodman, 1984], which is also described in [Bernstein et al, 1987], and is generally known as the *available copies protocol*. The coordinator of an update transaction T sends each $W_T(x)$ to all the sites where replicas of x reside and wait for confirmation of execution (or rejection). If it times out before it gets acknowledgement from all the sites, it considers those which have not replied as unavailable and continues with the update on the available sites. The unavailable sites, of course, update their databases to the latest state when they recover. Note, however, that these sites may not even be aware of the existence of T and the update to x that T has made if they had become unavailable before T started.

There are two complications that need to be addressed. The first one is the possibility that the sites that the coordinator thought were unavailable were in fact up and running and may have already updated x but their acknowledgement may not have reached the coordinator before its timer ran out. Second, some of these sites may have been unavailable when T started and may have recovered since then and have started executing transactions. Therefore, the coordinator undertakes a validation procedure before committing:

1. The coordinator checks to see if all the sites it thought were unavailable are still unavailable. It does this by sending an inquiry message to every one of these sites. Those that are available reply. If the coordinator gets a reply from one of these sites, it aborts T since it does not know the state that the previously unavailable site is in: it could have been that the site was available all along and had performed the original $W_T(x)$ but its acknowledgement was delayed (in which case everything is fine), or it could be that it was indeed unavailable when T started but became available later on and perhaps even executed $W_S(x)$ on behalf of another transaction S. In the latter case, continuing with T would make the execution schedule nonserializable.

2. If the coordinator of T does not get any response from any of the sites that it thought were unavailable, then it checks to make sure that all the sites that were available when $W_T(x)$ executed are still available. If they are, then T can proceed to commit. Naturally, this second step can be integrated into a commit protocol.

The ROWA-A class of protocols are more resilient to failures, including network partitioning, than the simple ROWA protocol. There have been a number of different variants of ROWA which we do not discuss further. Readers can consult [Helal et al., 1997] for further discussion.

Another class of strict replication protocols are those based on voting. The fundamental characteristics of voting were presented in the previous section when we discussed network partitioning in nonreplicated databases. The general ideas hold in replicated case. Fundamentally, each read and write operation has to obtain a sufficient number of votes to be able to commit. These protocols can be pessimistic or optimistic. In what follows we discuss only pessimistic protocols. An optimistic version [Davidson, 1984] compensates transactions to recover if the commit decision cannot be confirmed at completion. This version is suitable wherever compensating transactions are acceptable (see Chapter 10).

The initial voting algorithm was discussed in [Thomas, 1979] and an early suggestion to use quorum-based voting for replica control is due to [Gifford, 1979]. Thomas's algorithm works on fully replicated databases and assigns an equal vote to each site. For any operation of a transaction to execute, it must collect affirmative votes from a majority of the sites. Gifford's algorithm, on the other hand, works with partially replicated databases (as well as with fully replicated ones) and assigns a vote to each copy of a replicated data item. Each operation then has to obtain a *read quorum* (V_r) or a *write quorum* (V_w) to read or write a data item, respectively. If a given data item has a total of V votes, the quorums have to obey the following rules:

1. $V_r + V_w > V$
2. $V_w > V/2$

As the reader may recall from the preceding section, the first rule ensures that a data item is not read and written by two transactions concurrently (avoiding the read-write conflict). The second rule, on the other hand, ensures that two write operations from two transactions cannot occur concurrently on the same data item (avoiding write-write conflict). Thus the two rules ensure that serializability and one-copy equivalence are maintained.

In the case of network partitioning, the quorum-based protocols work well since they basically determine which transactions are going to terminate based on the votes that they can obtain. The vote allocation and threshold rules given above ensure that two transactions that are initiated in two different partitions and access the same data cannot terminate at the same time.

The difficulty with this version of the protocol is that transactions are required to obtain a quorum even to read data. This significantly and unnecessarily slows down read access to the database. We describe below another quorum-based voting protocol [Abbadi et al., 1985] that overcomes this serious performance drawback.

The protocol makes certain assumptions about the underlying communication layer and the occurrence of failures. The assumption about failures is that they are "clean." This means two things:

1. Failures that change the network's topology are detected by all sites instantaneously.
2. Each site has a view of the network consisting of all the sites with which it can communicate.

Based on the presence of a communication network that can ensure these two conditions, the replica control protocol is a simple implementation of the ROWA principle. When the replica control protocol attempts to read or write a data item, it first checks if a majority of the sites are in the same partition as the site at which the protocol is running. If so, it implements the ROWA rule within that partition: it reads any copy of the data item and writes all copies that are in that partition.

Notice that the read or the write operation will execute in only one partition. Therefore, this is a pessimistic protocol that guarantees one-copy serializability. When the partitioning is repaired, the database is recovered by propagating the results of the update to the other partitions.

The fundamental question with respect to implementation of this protocol is whether or not the failure assumptions are realistic. Unfortunately, they are not. Most network failures are not "clean." There is a time delay between the occurrence of a failure and its detection by a site. Because of this delay, it is possible for one site to think that it is in one partition when in fact subsequent failures have placed it in another partition. Furthermore, this delay may be different for various sites. Thus two sites that were in the same partition but are now in different partitions may proceed for a while under the assumption that they are still in the same partition. The violations of these two failure assumptions have significant negative consequences on the replica control protocol and its ability to maintain one-copy serializability.

The suggested solution is to build on top of the physical communication layer another layer of abstraction which hides the "unclean" failure characteristics of the physical communication layer and presents to the replica control protocol a communication service that has "clean" failure properties. This new layer of abstraction provides *virtual partitions* within which the replica control protocol operates. A virtual partition is a group of sites that have agreed on a common view of who is in that partition. Sites join and depart from virtual partitions under the control of this new communication layer, which ensures that the clean failure assumptions hold.

The advantage of this protocol is its simplicity. It does not incur any overhead to maintain a quorum for read accesses. Thus the reads can proceed as fast as they would in a nonpartitioned network. Furthermore, it is general enough so that the replica control protocol does not need to differentiate between site failures and network partitions.

There are many different versions of quorum-based protocols. Some of these are discussed in [Triantafillou and Taylor, 1995], [Paris, 1996], [Tanenbaum and van Renesse, 1988].

REVIEW QUESTIONS

12.1 Explain reliability concepts and measures.
12.2 What are the failures and fault tolerances in distributed system?
12.3 Explain the failures in DDBMS.
12.4 Explain by means of a diagram the interface between the local recovery manager and buffer manager.
12.5 Draw a diagram of logging interface.
12.6 Explain execution of LRM commands.
12.7 Draw a diagram for full memory hierarchy managed by LRM and BM.
12.8 Explain distributed reliability protocols.
12.9 Give an algorithm for 2PC coordinator.
12.10 How do you deal with site failure?
12.11 Draw a diagram of state transactions in 3PC protocol.
12.12 Draw a diagram of 3PC protocol actions.
12.13 Explain networking partitioning.