Data Bases II
Theory

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Abstract

The course aims to prepare software designers on the effective development of database applications.

First, the course presents the fundamental features of current database architectures, with a specific emphasis on the concept of transaction and its realization in centralized and distributed systems.

Then, the course illustrates the main directions in the evolution of database systems, presenting approaches that go beyond the relational model, like active databases, object systems and XML data management solutions.

Contents

Chapter 1

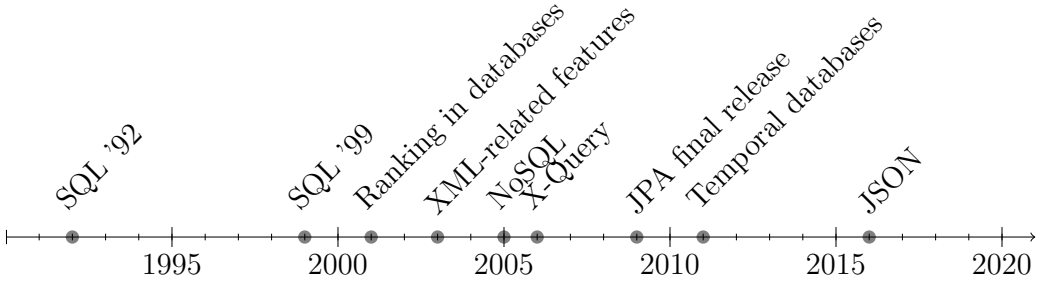
Introduction

1.1 Data Base Management System

Definition

A *Data Base Management System* is a software product capable of managing data collections that are:

- Large: much larger than the central memory available on the computers that run the software.
- Persistent: with a lifetime which is independent of single executions of the programs that access them.
- Shared: used by several applications at a time.
- Reliable: ensuring tolerance to hardware and software failures.
- Data ownership respectful: by disciplining and controlling accesses.



1.2 Transactions

Definition

A *transaction* is an elementary, atomic unit of work performed by an application. Each transaction is conceptually encapsulated within two commands:

- Begin transaction.
- End transaction.

Within a transaction, one of the commands below is executed to signal the end of the transaction: commit or rollback.

Definition

The *On-Line Transaction Processing* (OLTP) is a system that supports the execution of transactions on behalf of concurrent applications.

The application can run many transactions. So, the transactions are part of the application and not vice-versa. The transactions follow the ACID property:

1. Atomicity: a transaction is an indivisible unit of execution. This means that all the operations in the transaction are executed or none is executed. The time in which commit is executed marks the instant in which the transaction ends successfully: an error before should cause the rollback and an error after should not alter the transaction. The rollback of the work performed can be caused by a rollback statement or by the DBMS. In case of a rollback, the work performed must be undone, bringing the database to the state it had before the start of the transaction. It is the application's responsibility to decide whether an aborted transaction must be redone or not.

2. Consistency: A transaction must satisfy the database integrity constraints, that is if the initial state S_0 is consistent then the final state S_f is also consistent. This is not necessarily true for the intermediate states S_i.

3. Isolation: the execution of a transaction must be independent of the concurrent execution of other transactions.

4. Durability: the effect of a transaction that has successfully committed will last forever, independently of any system fault.

Property	Actions	Architectural element
Atomicity	Abort-rollback-restart	Query Manager
Consistency	Integrity checking	Integrity Control System
Isolation	Concurrency control	Concurrency Control System
Durability	Recovery management	Reliability Manager

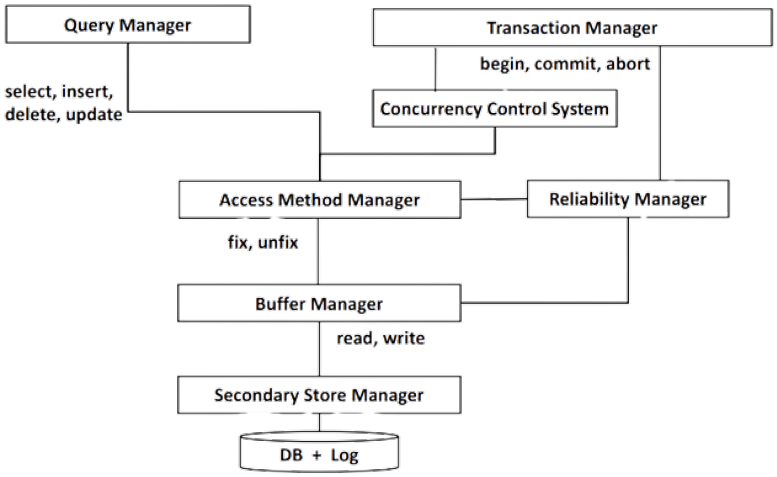


Figure 1.1: Architecture of a Data Base Management System

Chapter 2

Concurrency

2.1 Introduction

A DBMS usually needs to manage multiple applications. A unit of measurement used to evaluate the DBMS workload is the number of transaction per second (*tps*) handled by it. To have an efficient usage of the database the DBMS needs to be able to handle concurrency while avoiding the insurgence of anomalies. The concurrency control system schedules the order of the various transactions.

2.2 Anomalies in concurrent transactions

The anomalies caused by uncorrect scheduling are:

- Lost update: an update is applied from a state that ignores a preceding update, which is lost.

Transaction t_1	**Transaction** t_2
$r_1(x)$	
$x = x + 1$	
	$r_2(x)$
	$x = x + 1$
	$w_2(x)$
	commit
$w_1(x)$	
commit	

- Dirty read: an uncommitted value is used to update the data.

Transaction t_1	**Transaction** t_2
$r_1(x)$	
$x = x + 1$	
$w_1(x)$	
	$r_2(x)$
	commit
abort	

- Non-repeatable read: someone else updates a previously read value.

Transaction t_1	**Transaction** t_2
$r_1(x)$	
	$r_2(x)$
	$x = x + 1$
	$w_2(x)$
	commit
$r_1(x)$	
commit	

- Phantom update: someone else updates data that contributes to a previously valid constraint.

Transaction t_1	**Transaction** t_2
$r_1(x)$	
	$r_2(y)$
$r_1(y)$	
	$y = y - 100$
	$r_2(z)$
	$z = z + 100$
	$w_2(y)$
	$w_2(z)$
	commit
$r_1(z)$	
$s = x + y + z$	
commit	

- Phantom insert: someone else inserts data that contributes to a previously read datum.

2.3 Concurrency theory

Definition

A *model* is an abstraction of a system, object or process, which purposely disregards details to simplify the investigation of relevant properties.

Concurrency theory builds upon a model of transaction and concurrency control principles that help understanding the real systems. Real systems exploit implementation level mechanisms which help achieve some desirable properties postulated by the theory.

Definition

An *operation* consist in a reading or in a writing of a specific datum by a specific transaction.

A *schedule* is a sequence of operations performed by concurrent transactions that respects the order of operations of each transaction.

The transactions can be: serial, interleaved or nested. The number of serial schedules for n transaction is equal to:

$$N_S = n!$$

While the total number of distinct schedules given the number of transaction n is equal to:

$$N_D = \frac{(\sum_{i=1}^{n} k_i)!}{\prod_{i=1}^{n} (k_i!)}$$

Example : Given two transaction T_1 and T_2 we have six possible different schedules, where only two are serial:

1. $r_1(x)w_1(x)r_2(z)w2(z)$
2. $r_2(z)w_2(z)r_1(x)w_1(x)$
3. $r_1(x)r_2(z)w_1(x)w_2(z)$
4. $r_2(z)r_1(x)w_2(z)w_1(x)$
5. $r_1(x)r_2(z)w_2(z)w_1(x)$
6. $r_2(z)r_1(x)w_1(x)w_2(z)$

The first two are serial, the third and the fourth are nested, and the last two interleaved.

The concurrency control has to reject all the schedules that causes anomalies.

Definition

The *scheduler* is a component that accepts or rejects operations requested by the transactions.

The *serial schedule* is a schedule in which the actions of each transaction occur in a contiguous sequence.

A serializable schedule leaves the database in the same state as some serial schedule of the same transactions, so it is correct. To introduce other classes we have to initially make two assumptions:

- The transactions are observed a posteriori.
- Commit-projection: consider only the committed transactions.

2.4 View-serializability

Definition

$r_i(x)$ *reads-from* $w_j(x)$ when $w_j(x)$ precedes $r_i(x)$ and there is no $w_k(x)$ in S between $r_i(x)$ and $w_j(x)$.

$w_i(x)$ in a schedule S is a *final write* if it is the last write on x that occurs in S.

Two schedules are said to be *view-equivalent* ($S_i \approx_V S_j$) if they have:

1. The same operations.
2. The same reads-from relationships.
3. The same final writes.

A schedule is view-serializable (VSR) if it is view-equivalent to a serial schedule of the same transactions. The value written by $w_j(x)$ could be uncommitted when $r_i(x)$ reads it, but we are sure that it will be committed (commit-projection hypotesis).

Example : The following schedules are given:

- $S_1 : w_0(x)r_2(x)r_1(x)w_2(x)w_2(z)$
- $S_2 : w_0(x)r_1(x)r_2(x)w_2(x)w_2(z)$
- $S_3 : w_0(x)r_1(x)w_1(x)r_2(x)w_1(z)$
- $S_4 : w_0(x)r_1(x)w_1(x)w_1(z)r_2(x)$
- $S_5 : r_1(x)r_2(x)w_1(x)w_2(x)$

- $S_6 : r_1(x)r_2(x)w_2(x)r_1(x)$
- $S_7 : r_1(x)r_1(y)r_2(z)r_2(y)w_2(y)w_2(z)r_1(z)$

We have that only S_2 and S_3 are serial. S_1 is view-equivalent to serial schedule S_2 (so it is view-serializable). S_3 is not view-equivalent to S_2 (different operations) but is view-equivalent to serial schedule S_4, so it is also view-serializable.

S_5 corresponds to a lost update, S_6 corresponds to a non-repeatable read, and S_7 corresponds to a phantom update. All these schedules are non view-serializable.

The following schedules are given:

- $S_a : w_0(x)r_1(x)w_0(z)r_1(z)r_2(x)w_0(y)r_3(z)w_3(z)w_2(y)w_1(x)w_3(y)$
- $S_b : w_0(x)w_0(z)w_0(y)r_2(x)w_2(y)r_1(x)r_1(z)w_1(x)r_3(z)w_3(z)w_3(y)$
- $S_c : w_0(x)w_0(z)w_0(y)r_2(x)w_2(y)r_3(z)w_3(z)w_3(y)r_1(x)r_1(z)w_1(x)$

S_a and S_b are view-equivalent because all the reads-from relationship and final writes are the same. In fact, we have:

- Reads-from: $r_1(x)$ from $w_0(x)$, $r_1(z)$ from $w_0(z)$, $r_2(x)$ from $w_0(x)$, $r_3(z)$ from $w_0(z)$.
- Final writes: $w_1(x)$, $w_3(y)$, $w_3(z)$.

S_a and S_c are not view-equivalent because not all the reads-from relationship are the same.

Deciding if a generic schedule is in VSR is a NP-complete problem. Therefore, we need to find a stricter definition that is easier to check. The new definition may lead to rejecting some schedule that would be acceptable under view-serializability but not under the stricter criterion.

2.5 Conflict-serializability

Definition

Two operations o_i and o_j $(i \neq j)$ are in *conflict* if they address the same resource and at least one of them is write. There are two possible cases:

1. Read-write conflicts ($r - w$ or $w - r$).
2. Write-write conflicts ($w - w$).

Two schedules are *conflict-equivalent* ($S_i \approx_C S_j$) if S_i and S_j contain

the same operations and in all the conflicting pairs the transactions occur in the same order.

A schedule is *conflict-serializable* (CSR) if and only if it is conflict equivalent to a serial schedule of the same transactions.

We have that VSR is a strict subset of CSR and that CSR implies VSR.

Proof VSR is a subset of CSR : Schedule $S = r_1(x)w_2(x)w_1(x)w_3(x)$ is:

- View-serializable.
- Not conflict-serializable

It is possible to check that there is no conflict-equivalent serial schedule. ■

Proof CSR implies VSR : We assume $S_1 \approx_C S_2$ and prove that $S_1 \approx_V S_2$.
S_1 and S_2 must have:

- The same final writes: if they didn't, there would be at least two writes in a different order, and since two writes are conflicting operations, the schedules would not be $\approx_C$.
- The same reads-from relations: if not, there would be at least one pair of conflicting operations in a different order, and therefore, again, $\approx_C$ would be violated.

■

The testing of view-serializability is done with a conflict graph that has one node for each transaction T_i, and one arc from T_i to T_j if there exists at least one conflict between an operation o_i of T_i and an operation o_j of T_j such that o_i precedes o_j.

Theorem

A schedule is *conflict-serializable* if and only if its conflict graph is acyclic.

Example : We are given the schedule

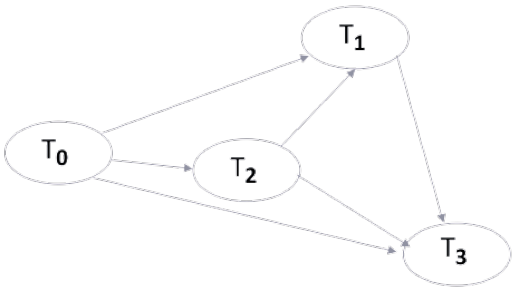
$$S : w_0(x)r_1(x)w_0(z)r_1(z)r_2(x)w_0(y)r_3(z)w_3(z)w_2(y)w_1(x)w_3(y)$$

To test the conflict serializability we have to do the following steps:

1. Create all the nodes based on the number of transactions of the schedule.
2. Divide the operation based on resource requested.
3. Check all the write-write and read-write relationships in each subset, and add the arcs based on these.

In the given example we obtain:

- $x : w_0 r_1 r_2 w_1$
- $y : w_0 w_2 w_3$
- $z : w_0 r_1 r_3 w_3$



There are no cycles in the graph, so this means that the schedule is CSR.

Proof CSR implies acyclicity of the conflict graph : Consider a schedule S in CSR. As such, it is $\approx_C$ to a serial schedule. Without loss of generality we can label the transactions of S to say that their order in the serial schedule is: $T_1 T_2 \dots T_n$. Since the serial schedule has all conflicting pairs in the same order as schedule S, in the conflict graph there can only be arcs (i, j), with $i < j$. Then the graph is acyclic, as a cycle requires at least an arc (i, j) with $i > j$. ■

Proof acyclicity of the conflict graph implies CSR : If S's graph is acyclic then it induces a topological (partial) ordering on its nodes. The same partial order exists on the transactions of S. Any serial schedule whose transactions are ordered according to the partial order is conflict-equivalent to S, because for all conflicting pairs (i, j) it is always $i < j$. ■

2.6 Concurrency control in practice

Conflict-serializability checking would be efficient if we knew the graph from the beginning, but usually we don't. Therefore, a scheduler must rather work online. So, it is not feasible to maintain the conflict graph, update it, and check its acyclicity at each operation request. At the same time, the assumption that concurrency control can work only with the commit-projection of the schedule is unrealistic because aborts do occur. We need

some simple online decision criterion for the scheduler, which must avoid as many anomalies as possible, and have negligible overhead.

When dealing with online concurrency control, it is important also to consider arrival sequences. The concurrency control system maps an arrival sequence into an effective a posteriori schedule. To implement this online scheduling we use two main families of techniques:

- Pessimistic (locks): if a resource is taken, make the requester wait or pre-empt the holder.
- Optimistic (timestamps and versions): serve as many requests as possible, possibly using out-of-date versions of the data.

Usually, commercial systems take the best of both worlds.

2.7 Locking

The method called locking is the most used in commercial systems. A transaction is well-formed with respect to locking if:

- Read operations are preceded by *r_lock* (shared) and followed by unlock.
- Write operations are preceded by *w_lock* (exclusive) and followed by unlock.

In both cases unlocking can be delayed with respect to the end of the operations. So, every object can be: free, r-locked or w-locked.

Transactions that first read and then write an object may acquire a *w_lock* already when reading or acquire a *r_lock* first and then upgrade it into a *w_lock* (escalation).

The lock manager receives requests from the transactions and grants resources according to the conflict table:

	Resource status		
Request	*FREE*	*R_LOCKED*	*W_LOCKED*
r_lock	✓ R_LOCKED	✓ R_LOCKED($n++$)	✗ W_LOCKED
w_lock	✓ W_LOCKED	✗ R_LOCKED	✗ W_LOCKED
unlock	ERROR	✓ $n--$	✓ FREE

Example : Given a schedule with three transactions and the following operations:

$$r_1(x)w_1(x)r_2(x)r_3(y)w_1(y)$$

We have the following locks:

- $r_1(x)$: $r_1\_lock(x)$ request $\rightarrow$ Ok $\rightarrow$ x is $r-locked$ with $n_x = 1$.
- $w_1(x)$: $w_1\_lock(x)$ request $\rightarrow$ Ok $\rightarrow$ x is $w-locked$.
- $r_2(x)$: $r_1\_lock(x)$ request $\rightarrow$ No, because x is $w-locked \rightarrow T_2$ waits.
- $r_3(y)$: $r_3\_lock(y)$ request $\rightarrow$ Ok $\rightarrow$ y is $r-locked$ with $n_y = 1$ and then T_3 unlocks y.
- $w_1(y)$: $w_1\_lock(y)$ request $\rightarrow$ Ok $\rightarrow$ y is $w-locked$ and then x and y are freed.

So, the schedule a posteriori will become:

$$r_1(x)w_1(x)r_3(y)w_1(y)r_2(x)$$

and we have that the transaction two is delayed.

The locking system is implemented via lock tables, which are hash tables indexing the lockable items via hashing and where each locked item has a linked list associated with it. Every node in the linked list represents the transaction which requested for lock, the lock mode and the current status. Every new lock request for the data item is appended to the list.

2.8 Two-phase locking

With the locking showed before we do not eliminate the anomalies caused by non-repeatable reads. To avoid this problem we can use a two-phase rule which requires that a transaction cannot acquire any other lock after releasing one. So, we have a phase where the transaction acquires all the locks, a phase where it executes all operations and a final phase of unlocking.

Definition

The class of *two-phase locking* is the set of all schedules generated by a scheduler that:

- Only processes well-formed transactions.
- Grant locks according to the conflict table.
- Checks that all transactions apply the two-phase rule.

We have that $2PL$ is a strict subset of CSR and also that $2PL$ implies CSR.

Proof 2PL implies CSR : We assume that a schedule S is $2PL$. Consider, for each transaction, the moment in which it holds all locks and is going to release the first one. We sort the transactions by this temporal value and consider the corresponding serial schedule $S^{'}$. We want to prove by contradiction that S is conflict-equivalent to $S^{'}$:

$$S^{'} \approx_C S, \ldots$$

Consider a generic conflict $o_i \rightarrow o_j$ in $S^{'}$ with $o_i \in T_i$, $0_j \in T_j$, and $i < j$. By definition of conflict, o_i and o_j address the same resource r, and at least one of them is write. The two operations cannot occur in reverse order of S. This proves that all $2PL$ schedules are view-serializable. ■

The anomalies that remain in this state are only the phantom inserts (needs predicate locks) and the dirty reads.

2.9 Strict two-phase and predicate locking

Definition

In *strict two-phase locking* (or long duration locks) we also have that locks held by a transaction can be released only after commit or rollback.

This version of locking is used in most commercial DBMS whenever a high level of isolation is required.

To prevent phantom inserts a lock should be placed also on future data using the predicate locks. If the predicate lock is on a resource, other transactions cannot insert, delete, or update any tuple satisfying this predicate.

2.10 Isolation levels in SQL '99

SQL defines transaction isolation levels which specify the anomalies that should be prevented by running at that level:

	Dirty read	**Non-repeatable read**	**Phantoms**
Read uncommitted	✓	✓	✓
Read committed	✗	✓	✓
Repeatable reads	✗	✗	✓(insert)
Serializable	✗	✗	✗

The four levels are implemented respectively with: no read locks, normal read locks, strict read locks, and strict locks with predicate locks. Serializable is not the default because its strictness can arise the following problems:

- Deadlock: two or more transactions in endless mutual wait.
- Starvation: a single transaction in endless wait.

2.11 Deadlocks

A deadlock occurs because concurrent transactions hold and in turn request resources held by other transactions.

Definition

A *lock graph* is a bipartite graph in which nodes are resources or transactions and arcs are lock requests or lock assignments.

A *wait-for graph* is a graph in which nodes are transactions and arcs are waits for relationships.

A deadlock is represented by a cycle in the wait-for graph of transactions.

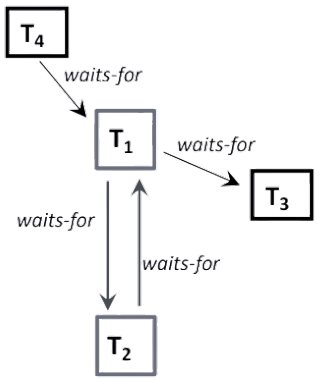


Figure 2.1: An example of deadlock in the wait-for graph

It is possible to solve deadlocks in three different ways:

- Timeout: a transaction is killed and restarted after a given amount of waiting, to be determined by the system manager.
- Deadlock prevention: kills transactions that could cause cycles. It is implemented in two ways:

1. Resource-based prevention puts restrictions on lock requests. The idea is that every transaction requests all resources at once, and only once. The main problem is that it's not easy for transactions to anticipate all requests.
2. Transaction-based prevention puts restrictions on transactions' IDs. Assigning IDs to transactions incrementally allows to give an age to each one. It is possible to choose to kill the holding transaction (preemptive) or the requesting one (non-preemptive). The main problem is that the number of killings is too big.

- Deadlock detection: it can be implemented with various algorithms and used for distributed resources.

Definition

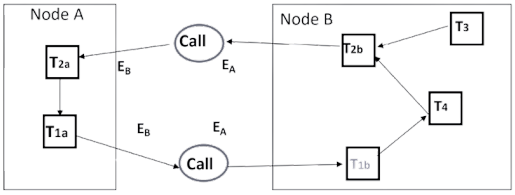
The *distributed dependency graph* is a wait-for graph where external call nodes represent a sub-transaction activating another sub-transaction at a different node.

The arrow shows a wait-for relation among local transactions. If one term is an external call, either the source is waited for by a remote transaction or waits for a remote transaction. The Obermarck's algorithm needs the following assumptions:

- Transactions execute on a single main node.
- Transactions may be decomposed in sub-transactions running on other nodes.
- When a transaction spawns a sub-transaction it suspends work until the latter completes.
- Two wait-for relationships:
  - T_i waits for T_j on the same node because T_i needs a datum locked by T_j.
  - A sub-transaction of T_i waits for another sub-transaction of T_i running on a different node.

The goal of this algorithm is to detect a potential deadlock looking only at the local view of a node. Nodes exchange information and update their local graph based on the received information. Node A sends its local info to a node B only if: it contains a transaction T_i that is waited for from another remote transaction and waits for a transaction T_j active on B and $i > j$.

Example : Given the following distributed dependency graph:



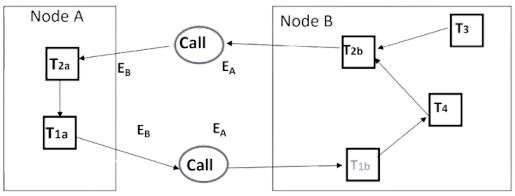
We can see that the potential deadlock is given by cycles. In this case, we have that T_{2a} waits for T_{1a} (data lock) that waits for T_{1b} (call) that waits for T_{2b} (data locks) that waits for T_{2a} (call).

In this case the node A dispatches information to B, in fact we have $E_b \rightarrow T_2 \rightarrow T_1 \rightarrow E_b$ and node B cannot dispatch information to A, because the forwarding rule is not respected: $E_a \rightarrow T_1 \rightarrow T_2 \rightarrow E_a$.

The Obermarck's algorithm runs periodically at each node and consists in four steps:

1. Get graph info from the previous nodes.
2. Update the local graph by merging the received information.
3. Check the existence of cycles among transactions denoting potential deadlocks: if found, select one transaction in the cycle and kill it.
4. Send updated graph info to the next nodes.

Example : Given the following distributed system:



We want to apply the Obermarck's algorithm:

1. Use the forwarding rule, in this case we have:
   - At Node A: $E_b \rightarrow T_2 \rightarrow T_1 \rightarrow E_b$ info sent to Node B
   - At Node B: $E_a \rightarrow T_1 \rightarrow T_2 \rightarrow E_a$ info not sent $(i < j)$.
2. At node B there is the updated info $E_b \rightarrow T_2 \rightarrow T_1 \rightarrow E_b$ and it is added to the wait-for graph.

3. At node B a deadlock is detected (cycle between T_1 and T_2) and T_1, T_2 or T_4 are killed.
4. Updated information are sent to all nodes.

There are four variants of this algorithm, based on various conventions.

	Message condition	**Message receiver**
A	$i > j$	Following node
B	$i > j$	Preceding node
C	$i < j$	Following node
D	$i < j$	Preceding node

In practice, the probability of deadlocks (n^{-2}) is much less than the conflict probability (n^{-1}). There are techniques to limit the frequency of deadlocks:

- Update lock: the most frequent deadlock occurs when two concurrent transactions start by reading the same resources and then decide to write and try to upgrade their lock to write on the resource. To avoid this situation, systems offer the update lock, that is used by transactions that will read and then write. The lock table become:

	Resource status			
Request	*FREE*	*SHARED*	*UPDATE*	*EXCLUSIVE*
Shared lock	✓	✓	✓	✗
Update lock	✓	✓	✗	✗
Exclusive lock	✓	✗	✗	✗

- Hierarchical lock: locks can be specified with different granularities. The objective of this is to lock the minimum amount of data and recognize conflicts as soon as possible. The method used to do so consists in asking locks on hierarchical resources by requesting resources top-down until the right level is obtained and releasing locks bottom-up. This is done by using five locking modes: shared, exclusive, ISL (intention of locking a sub-element of the current element in shared mode), IXL (intention of locking a sub-element of the current element in exclusive mode), and SIXL (lock of the element in shared mode with intention of locking a sub-element in exclusive mode). The lock table become:

	Resource status					
Request	*FREE*	*ISL*	*IXL*	*SL*	*SIXL*	*XL*
ISL	✓	✓	✓	✓	✓	✗
IXL	✓	✓	✓	✗	✗	✗
SL	✓	✓	✗	✓	✗	✗
SIXL	✓	✓	✗	✗	✗	✗
XL	✓	✗	✗	✗	✗	✗

Example : Given a table X with eight tuples divided in two pages:

P1	**P2**
$t1$	$t5$
$t2$	$t6$
$t3$	$t7$
$t4$	$t8$

And two transactions with the following schedules:

$$T_1 = r(P1)\ w(t3)\ r(t8)$$

$$T_2 = r(t2)\ r(t4)\ w(t5)\ w(t6)$$

We can see that they are not in a read-write conflict (because they are independent of the order). Without hierarchical locking both transactions needs to operate on the same table, so the concurrency will be almost useless in this case. But with this technique, calling X the table, we have that the transaction acquires the following locks:

$$T_1 :\ IXL(root)\ SIXL(P1)\ XL(t3)\ ISL(P2)\ SL(t8)$$

$$T_2 :\ IXL(root)\ ISL(P1)\ SL(t2)\ SL(t4)\ IXL(P2)\ XL(t5)\ XL(t6)$$

2.12 Timestamps

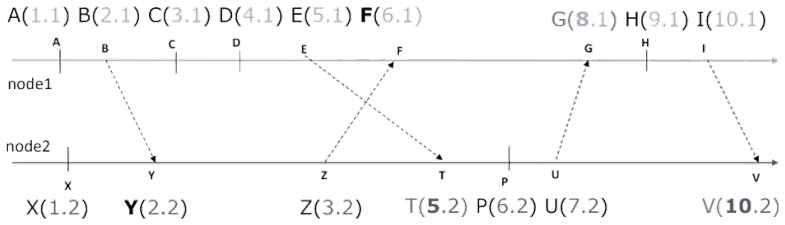
Locking is also named pessimistic concurrency control because it assumes that collisions will arise. In reality collisions are rare. So, it is possible to use optimistic concurrency control methods like timestamps, which are identifier that defines a total ordering of the events of a system. Each transaction has a timestamp representing the time at which the transaction begins so that transactions can be ordered: the smaller is the index the older is the transaction. A schedule is accepted only if it reflects the serial ordering of

the transactions induced by their timestamps. The timestamps are given by a system's function on request. The syntax of a timestamp is the following:

event-id.node-id

The algorithm's synchronization is based on send-receive of messages, and it is called Lamport method. It is based on the following rule: it is not possible to receive a message from the future, if this happens the bumping rule is used to bump the timestamp of the receive event beyond the timestamp of the send event.

Example: An example of timestamps assignation at two different nodes can be the following.



The scheduler uses two counters: one for writes ($WTM(x)$) and another for reads ($RTM(x)$). The scheduler receives read/write requests tagged with the timestamp of the requesting transaction. In case of a read operation we can have:

- If $ts < WTM(x)$ the request is rejected and the transaction is killed.
- Else, access is granted, and we set $RTM(x) = \max(RTM(x), ts)$.

In case of a write operation we can have:

- If $ts < RTM(x)$ or $ts < WTM(x)$ the request is rejected, and the transaction is killed.
- Else, access is granted, and we set $WTM(x) = ts$.

The problem that arise using this technique is that too many transactions are killed.

Example: Let us assume $RTM(x) = 7$ and $WTM(x) = 4$ and the following schedule:

$$S = r_6(x)r_8(x)r_9(x)w_8(x)w_{11}(x)r_{10}(x)$$

Using the timestamps we obtain:

Request	Response	New value
$r_6(x)$	✓	-
$r_8(x)$	✓	$RTM(x) = 8$
$r_9(x)$	✓	$RTM(x) = 9$
$w_8(x)$	✗	T_8 killed
$w_{11}(x)$	✓	$WTM(x) = 11$
$r_{10}(x)$	✗	T_{10} killed

It is not possible to compare 2PL to TS (this means that there is no subset between the categories). We have only that TS implies CSR.

Proof TS implies CSR : Let S be a TS schedule of T_1 and T_2. Suppose S is not CSR, which implies that it contains a cycle between T_1 and T_2. S contains $op_1(x)$, $op_2(x)$ where at least one is a write. S contains also $op_2(y)$, $op_1(y)$ where at least one is a write. When $op_1(y)$ arrives:

- If $op_1(y)$ is a read, T_1 is killed by TS because it tries to read a value written by a younger transaction, so it is a contradiction.
- If $op_1(y)$ is a write, T_1 is killed no matter what $op_2(y)$ is, because it tries to write a value already read or written by a younger transaction, so it is a contradiction.

■

Basic TS-based control considers only committed transactions in the schedule, aborted transactions are not considered. If aborts occur, dirty reads may happen. To cope with dirty reads, a variant of basic TS must be used: a transaction T_i that issues $r_{ts}(x)$ or $w_{ts}(x)$ such that $ts > WTM(x)$ has its read or write operation delayed until the transaction T' that wrote the value of x has committed or aborted. This is similar to long duration write locks, so it introduces delays.

Action	2PL	TS
Transaction	Wait	Killed and restarted
Serialization	Imposed by conflicts	Imposed by timestamp
Delay	Long (strict version)	Long
Deadlocks	Possible	Prevented

Since that restarting a transaction is costlier than waiting, 2PL is better if used alone. Commercial system mediates between those techniques to get the better feature from both. To reduce the number of killings it is possible to use the Thomas rule, that changes the rule for the write operations:

- If $ts < RTM(x)$ the request is rejected and the transaction is killed.
- Else, if $ts < WTM(x)$ then our write is obsolete: it can be skipped.
- Else, access is granted, and we set $WTM(x) = ts$.

2.13 Multi-version concurrency control

The idea of multi-version is that writes generate new versions, and that reads access the right version. Writes generate new copies, each one with a new WTM, so each object x always has $N \geq 1$ active versions. There is a unique global $RTM(x)$. Old versions are discarded when there are no transactions that need their values.

In theory, it is possible to use the following rule:

- $r_{ts}(x)$ is always accepted. A copy x_k is selected for reading such that:
  - If $ts \geq WTM_N(x)$, then $k = N$.
  - Else take k such that $WTM_k(x) \leq ts < WTM_{k+1}(x)$.
- $w_{ts}(x)$:
  - If $ts < RTM(x)$ the request is rejected.
  - Else a new version is created for timestamp ts (N is incremented). $WTM_1(x), \ldots, WTM_N(x)$ are the new versions, kept sorted from oldest to youngest.

Example: Let us assume $RTM(x) = 7$, $N = 1$ and $WTM_1(x) = 4$ and the following schedule:

$$S = r_6(x) r_8(x) r_9(x) w_8(x) w_{11}(x) r_{10}(x) r_{12}(x) w_{14}(x) w_{13}(x)$$

Using the multi-version we obtain:

Request	**Response**	**New value**
$r_6(x)$	✓	-
$r_8(x)$	✓	$RTM(x) = 8$
$r_9(x)$	✓	$RTM(x) = 9$
$w_8(x)$	✗	T_8 killed
$w_{11}(x)$	✓	$WTM_2(x) = 11,\ N = 2$
$r_{10}(x)$	✓ on $x_{(1)}$	$RTM(x) = 10$
$r_{12}(x)$	✓ on $x_{(2)}$	$RTM(x) = 12$
$w_{14}(x)$	✓	$WTM_3(x) = 14,\ N = 3$
$w_{13}(x)$	✓	$WTM_4(x) = 14,\ N = 4$

In practice, it is possible to use the following rule:

- $r_{ts}(x)$ is always accepted. A copy x_k is selected for reading such that:
  - If $ts \geq WTM_N(x)$, then $k = N$.
  - Else take k such that $WTM_k(x) \leq ts < WTM_{k+1}(x)$.
- $w_{ts}(x)$:
  - If $ts < RTM(x)$ or $ts < WTM_N(x)$ the request is rejected.
  - Else a new version is created for timestamp ts (N is incremented)
  - $WTM_1(x), \ldots, WTM_N(x)$ are the new versions, kept sorted from oldest to youngest.

Example : Let us assume $RTM(x) = 7$, $N = 1$ and $WTM_1(x) = 4$ and the following schedule:

$$S = r_6(x)r_8(x)r_9(x)w_8(x)w_{11}(x)r_{10}(x)r_{12}(x)w_{14}(x)w_{13}(x)$$

Using the multi-version we obtain:

Request	**Response**	**New value**
$r_6(x)$	✓	-
$r_8(x)$	✓	$RTM(x) = 8$
$r_9(x)$	✓	$RTM(x) = 9$
$w_8(x)$	✗	T_8 killed
$w_{11}(x)$	✓	$WTM_2(x) = 11,\ N = 2$
$r_{10}(x)$	✓ on $x_{(1)}$	$RTM(x) = 10$
$r_{12}(x)$	✓ on $x_{(2)}$	$RTM(x) = 12$
$w_{14}(x)$	✓	$WTM_3(x) = 14,\ N = 3$
$w_{13}(x)$	✗	T_{13} killed

The final structure for the sets is the following.

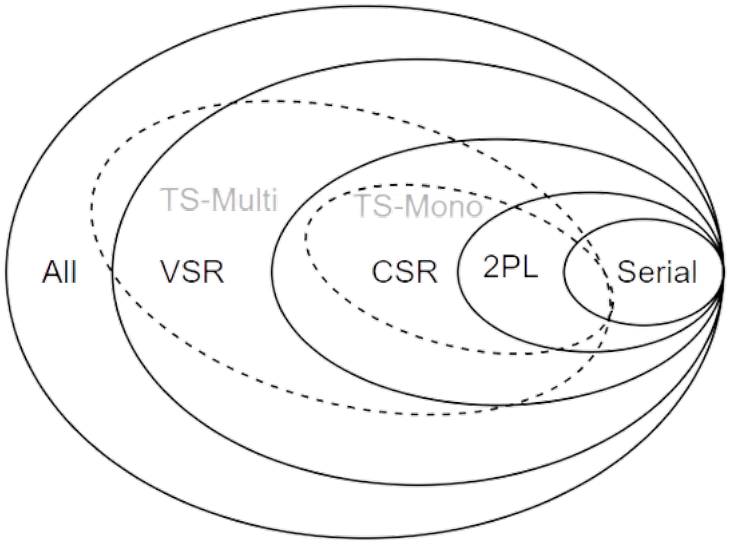


Figure 2.2: Set structure for concurrency classification

The realization of TS-Multi gives the opportunity to introduce into the DBMS another isolation level called snapshot isolation. In this level only $WTM(x)$ is used. The rule used in this level states that every transaction reads the version consistent with its timestamp, and defers writes to the end. If the scheduler detects that the writes of a transaction conflict with writes of other concurrent transactions after the snapshot timestamp, it aborts. Snapshot isolation does not guarantee serializability: it is possible to have a new anomaly called write skew (the order of the operation can change in different executions changing the final result).

Chapter 3

Ranking