Data Bases II
Theory

Christian Rossi

Academic Year 2023-2024

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

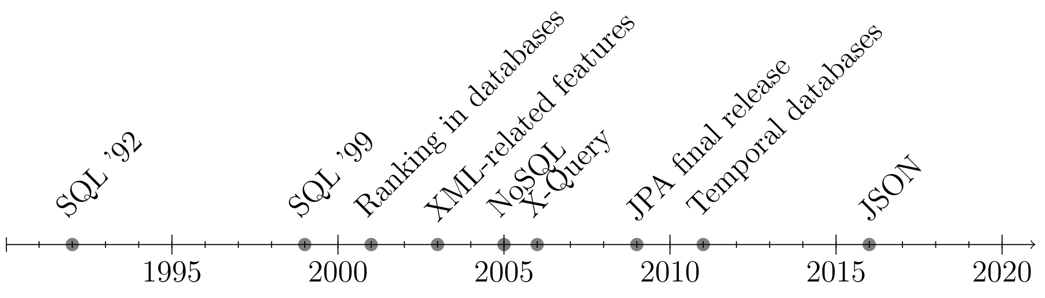
Introduction

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 and 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: 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. In particular, the concurrent execution of a number of transactions must produce the same result as the execution of the same transactions in a sequence.

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

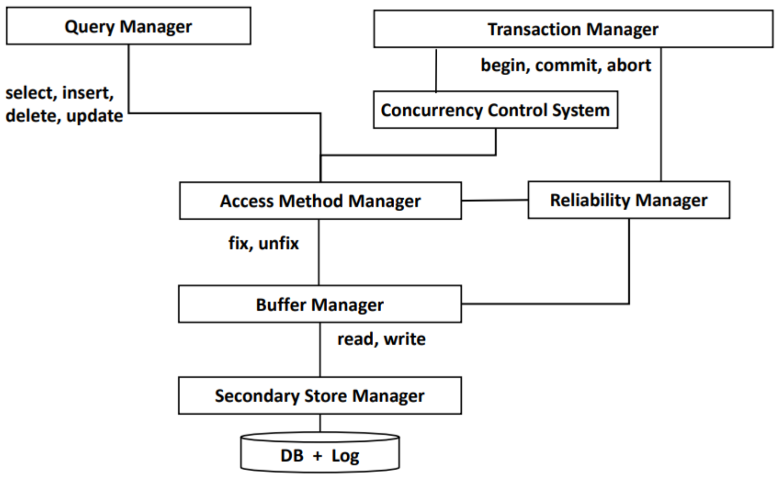


Figure 1.1: Architecture of a Data Base Management System

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.

Definition

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.
- The transactions limited to those that have been committed (commit-projection).

2.4 View-serializability

Definition

We say that:

- $r_i(x)$ *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 reads-from relationship $r_i(x)$ reads-from $w_j(x)$ in a schedule S assumes that $r_i(x)$ reads the value written by $w_j(x)$ independently of the time at which the commit of T_j occurs. In other words the value written by $w_j(x)$ could be uncommitted when $r_i(x)$ reads it, but we are sure (by definition of commit-projection) that it will be committed.

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/faster 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 a write. There are tho 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 $VSR \subset CSR$ 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 in conflict-serializable if and only if its conflict graph is acyclic.

Example : We are given the schedule

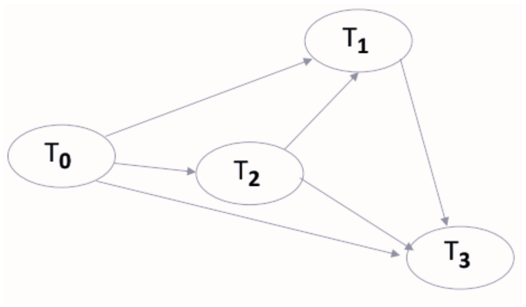
$$S : w_0(x)r_1(x)w_0(z)r_1(z)r_2(x)w_0(y)r_3(z)w_3(z)w_2(y)w_1(x)w_3(y)$$

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:

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



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 on-line 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 on-line 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 read/write operation. 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 became:

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

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 ocked item has a linked list associated with it. Every node in the linked list represents the transaction which requested for lock, lock mode and current status. Every new lock request for the data item is appended as a new node 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 a lock. So, we have a phase where the transaction acquires all the locks, a phase where it executes all operations and a fianl 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.
- Grants locks according to the conflict table.
- Checks that all transactions apply the two-phase rule.

We have that $2PL \subset CSR \subset VSR$.

Proof 2PL implies CSR : We assume that a schedule S is $2PL$. Consider, for each transaction, the end of the plateau (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 $S^{'}$. In order to prove (by contradiction) that S is conflict-equivalent to $S^{'}, S^{'} \approx_C S, \ldots$. Consider a generic conflict $o_i \to o_j$ 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 a write. The two operations cannot occur in reverse oreder of S. This proves that all 2PL schedules are view-serializable too and they can be checked with negligible overhead. ■

The anomalies that remains in this state are only the phantom insert (solvable with predicate locks) and the dirty read (requires dealing with abort).

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 (mainly for read operations) which specify the anomalies that should be prevented by running at that level:

- Read uncommited: allows dirty reads, non-repeatable reads and phantom updates and inserts. No read locks are used.
- Read committed: prevents dirty reads but allows non-repeatable reads and phantom updates/inserts. Normal read locks are used.
- Repeatable read: avoids dirty reads, non-repeatable reads and phantom updates, but allows phantom inserts. Strict read locks are used.

- Serializable: avoids all anomalies. Strict 2PL with predicate locks are used.

	Dirty read	Non-repeatable read	Phantoms
Read uncommitted	✓	✓	✓
Read committed	×	✓	✓
Repeatable reas	×	×	✓(insert)
Serializable	×	×	×

The strictness of serializable level can lead to deadlocks. Therefore, it is not the default in most commercial systems. Transactions requesting locks are either granted the lock or suspended and queued. There is risk of:

- 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.

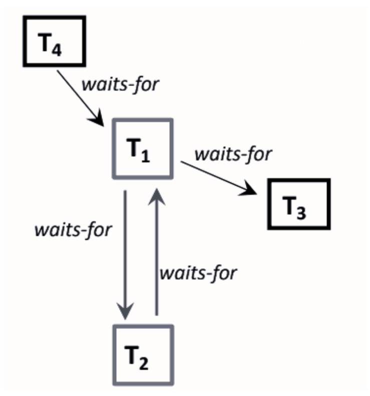


Figure 2.1: An example of deadlock in the wait-for graph

It is possible to solve deadlocks in three differemt ways:

- Timeout: a transaction is killed and restarted after a given amount of waiting. The timeout value is system-determined. The problem is choosing a proper timeout value.
- 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 the number of killings is too big.
- Deadlock detection: it can be implemented with Obermarck's algorithm 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.

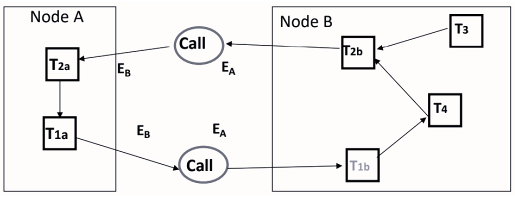
The arrow shows a wait-for relation among local transactions. If one term is an external call, either the source is waited for by a remote transaction or the sink waits for a remote transaction. The Obermarck's algorithm needs the following assumptions:

- 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 using a communication protocol whereby each node has a local projection of the global dependencies. 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 $i > j$. In other words, I send info to you if a distributed transaction listed at me waits for a distributed transaction listed at you with smaller index.

Example : Give the following distributed dependency graph:



We can see that the potential deadlock is given by the cycle. In fact, 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, becasue 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.