Week 09

Transaction Processing

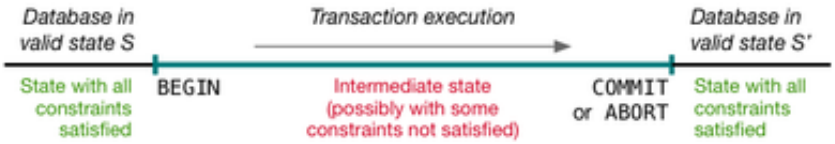
Transaction Processing

2/103

A *transaction* (tx) is ...

- a single application-level operation
- performed by a sequence of database operations

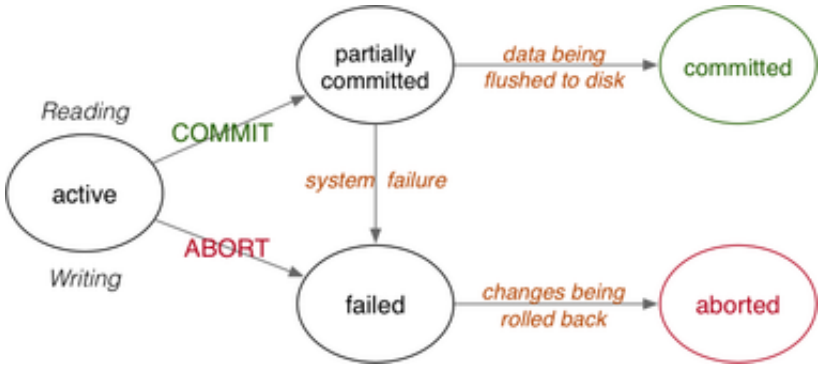
A transaction effects a state change on the DB



... Transaction Processing

3/103

Transaction states:



`COMMIT` ⇒ all changes preserved, `ABORT` ⇒ database unchanged

... Transaction Processing

4/103

Concurrent transactions are

- desirable, for improved performance (throughput)
- problematic, because of potential unwanted interactions

To ensure problem-free concurrent transactions:

- **A**tomic ... whole effect of tx, or nothing
- **C**onsistent ... individual tx's are "correct" (wrt application)
- **I**solated ... each tx behaves as if no concurrency
- **D**urable ... effects of committed tx's persist

... Transaction Processing

5/103

Transaction processing:

- the study of techniques for realising ACID properties

Consistency is the property:

- a tx is correct with respect to its own specification
- a tx performs a mapping that maintains all DB constraints

Ensuring this must be left to application programmers.

Our discussion focusses on: Atomicity, Durability, Isolation

... Transaction Processing

6/103

Atomicity is handled by the *commit* and *abort* mechanisms

- **commit** ends tx and ensures all changes are saved
- **abort** ends tx and *undoes* changes "already made"

Durability is handled by implementing *stable storage*, via

- redundancy, to deal with hardware failures
- logging/checkpoint mechanisms, to recover state

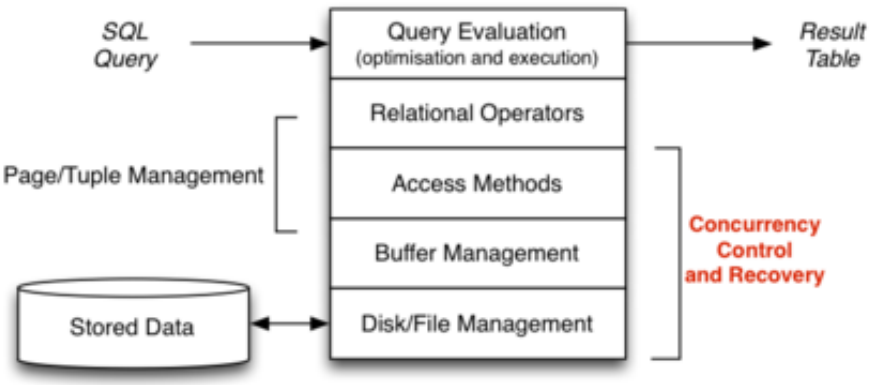
Isolation is handled by *concurrency control* mechanisms

- possibilities: lock-based, timestamp-based, check-based
- various levels of isolation are possible (e.g. serializable)

... Transaction Processing

7/103

Where transaction processing fits in the DBMS:



Transaction Terminology

8/103

To describe transaction effects, we consider:

- `READ` - transfer data from "disk" to memory
- `WRITE` - transfer data from memory to "disk"
- `ABORT` - terminate transaction, unsuccessfully
- `COMMIT` - terminate transaction, successfully

Relationship between the above operations and SQL:

- **`SELECT`** produces `READ` operations on the database
- **`UPDATE`** and **`DELETE`** produce `READ` then `WRITE` operations
- **`INSERT`** produces `WRITE` operations

... Transaction Terminology

9/103

More on transactions and SQL

- **`BEGIN`** starts a transaction
  - the `begin` keyword in PLpgSQL is not the same thing
- **`COMMIT`** commits and ends the current transaction
  - some DBMSs e.g. PostgreSQL also provide `END` as a synonym
  - the `end` keyword in PLpgSQL is not the same thing
- **`ROLLBACK`** aborts the current transaction, undoing any changes
  - some DBMSs e.g. PostgreSQL also provide `ABORT` as a synonym

In PostgreSQL, tx's cannot be defined inside functions (e.g. PLpgSQL)

... Transaction Terminology

10/103

The `READ`, `WRITE`, `ABORT`, `COMMIT` operations:

- occur in the context of some transaction *T*
- involve manipulation of data items *X, Y, ...* (`READ` and `WRITE`)

The operations are typically denoted as:

$R_T(X)$	read item *X* in transaction *T*
$W_T(X)$	write item *X* in transaction *T*
A_T	abort transaction *T*
C_T	commit transaction *T*

Schedules

11/103

A *schedule* gives the sequence of operations from *≥ 1* tx

Serial schedule for a set of tx's T_1 .. T_n

- all operations of T_i complete before T_{i+1} begins

E.g. $R_{T_1}(A)$ $W_{T_1}(A)$ $R_{T_2}(B)$ $R_{T_2}(A)$ $W_{T_3}(C)$ $W_{T_3}(B)$

Concurrent schedule for a set of tx's T_1 .. T_n

- operations from individual T_i's are interleaved

E.g. $R_{T_1}(A)$ $R_{T_2}(B)$ $W_{T_1}(A)$ $W_{T_3}(C)$ $W_{T_3}(B)$ $R_{T_2}(A)$

... Schedules

12/103

Serial schedules guarantee database consistency

- each T_i commits before T_{i+1}
- prior to T_i database is consistent
- after T_i database is consistent (assuming T_i is correct)
- before T_{i+1} database is consistent ...

Concurrent schedules interleave tx operations arbitrarily

- and may produce a database that is not consistent
- after all of the transactions have committed successfully

Transaction Anomalies

13/103

What problems can occur with (uncontrolled) concurrent tx's?

The set of phenomena can be characterised broadly under:

- *dirty read*:
  reading data item written by a concurrent uncommitted tx
- *nonrepeateable read*:
  re-reading data item, since changed by another concurrent tx
- *phantom read*:
  re-scanning result set, finding it changed by another tx

Exercise 1: Update Anomaly

14/103

Consider the following transaction (expressed in pseudo-code):

```
-- Accounts(id,owner,balance,...)
transfer(src id, dest id, amount int)
{
   -- R(X)
   select balance from Accounts where id = src;
   if (balance >= amount) {
      -- R(X),W(X)
      update Accounts set balance = balance-amount
      where id = src;
      -- R(Y),W(Y)
      update Accounts set balance = balance+amount
      where id = dest;
}  }
```

If two transfers occur on this account simultaneously,
give a schedule that illustrates the "dirty read" phenomenon.

Exercise 2: How many Schedules?

15/103

In the previous exercise, we looked at several schedules

For a given set of tx's T_1 ... T_n ...

- how many serial schedules are there?
- how many total schedules are there?

Schedule Properties

16/103

If a concurrent schedule on a set of tx's *TT* ...

- produces the same effect as a serial schedule on *TT*
- then we say that the schedule is *serializable*

Primary goal of isolation mechanisms (see later) is

- arrange execution of individual operations in tx's in *TT*
- to ensure that a serializable schedule is produced

Serializability is one property of a schedule, focusing on isolation

Other properties of schedules focus on recovering from failures

---

Transaction Failure

17/103

So far, have implicitly assumed that all transactions commit.

Additional problems can arise when transactions abort.

Consider the following schedule where transaction T1 fails:

```
T1: R(X) W(X) A
T2:             R(X) W(X) C
```

Abort *will* rollback the changes to `X`, but ...

Consider three places where the rollback might occur:

```
T1: R(X) W(X) A [1]     [2]         [3]
T2:                 R(X)    W(X) C
```

... Transaction Failure

18/103

Abort / rollback scenarios:

```
T1: R(X) W(X) A [1]     [2]         [3]
T2:                 R(X)    W(X) C
```

Case [1] is ok

- all effects of T1 vanish; final effect is simply from T2

Case [2] is problematic

- some of T1's effects persist, even though T1 aborted

Case [3] is also problematic

- T2's effects are lost, even though T2 committed

Recoverability

19/103

Consider the serializable schedule:

```
T1:         R(X)   W(Y)   C
T2:   W(X)                     A
```

(where the final value of `Y` is dependent on the `X` value)

Notes:

- the final value of X is valid (change from T_2 rolled back)
- T_1 reads/uses an X value that is eventually rolled-back
- even though T_2 is correctly aborted, it has produced an effect

Produces an invalid database state, even though serializable.

... Recoverability

20/103

Recoverable schedules avoid these kinds of problems.

For a schedule to be recoverable, we require additional constraints

- all tx's T_i that wrote values used by T_j
- must have committed before T_j commits

and this property must hold for all transactions T_j

Note that recoverability does not prevent "dirty reads".

In order to make schedules recoverable in the presence of dirty reads and aborts, may need to abort multiple transactions.

Exercise 3: Recoverability/Serializability

21/103

Recoverability and **S**erializability are orthogonal, i.e.

- a schedule can be R & S, !R & S, R &!S, !R & !S

Consider the two transactions:

```
T1:   W(A)   W(B)   C
T2:   W(A)   R(B)   C
```

Give examples of schedules on T1 and T2 that are

- recoverable and serializable
- not recoverable and serializable
- recoverable and not serializable

Cascading Aborts

22/103

Recall the earlier non-recoverable schedule:

```
T1:         R(X)   W(Y)   C
```

```
T2:  W(X)                  A
```

To make it recoverable requires:

- delaying T_1's commit until T_2 commits
- if T_2 aborts, cannot allow T_1 to commit

```
T1:        R(X)  W(Y) ...   C? A!
T2:  W(X)                  A
```

Known as *cascading aborts* (or *cascading rollback*).

... Cascading Aborts

23/103

Example: T_3 aborts, causing T_2 to abort, causing T_1 to abort

```
T1:                      R(Y)  W(Z)         A
T2:        R(X)  W(Y)                  A
T3:  W(X)                          A
```

Even though T_1 has no direct connection with T_3
(i.e. no shared data).

This kind of problem ...

- can potentially affect very many concurrent transactions
- could have a significant impact on system throughput

... Cascading Aborts

24/103

Cascading aborts can be avoided if

- transactions can only read values written by committed transactions

  (alternative formulation: no tx can read data items written by an uncommitted tx)

Effectively: eliminate the possibility of reading dirty data.

Downside: reduces opportunity for concurrency.

GUW call these *ACR* (avoid cascading rollback) schedules.

All ACR schedules are also recoverable.

Strictness

25/103

Strict schedules also eliminate the chance of *writing* dirty data.

A schedule is *strict* if

- no tx can read values written by another uncommitted tx (ACR)
- no tx can write a data item written by another uncommitted tx

Strict schedules simplify the task of rolling back after aborts.

... Strictness

26/103

Example: non-strict schedule

```
T1:  W(X)        A
T2:       W(X)     A
```

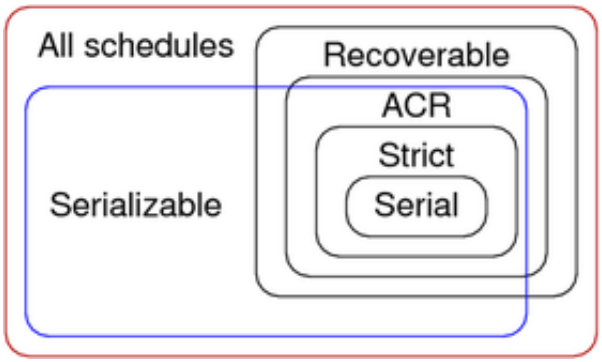
Problems with handling rollback after aborts:

- when T_1 aborts, don't rollback (need to retain value written by T_2)
- when T_2 aborts, need to rollback to pre-T_1 (not just pre-T_2)

Classes of Schedules

27/103

Relationship between various classes of schedules:



Schedules ought to be serializable and strict.

But more serializable/strict ⇒ less concurrency.

DBMSs allow users to trade off "safety" against performance.

Transaction Isolation

Transaction Isolation

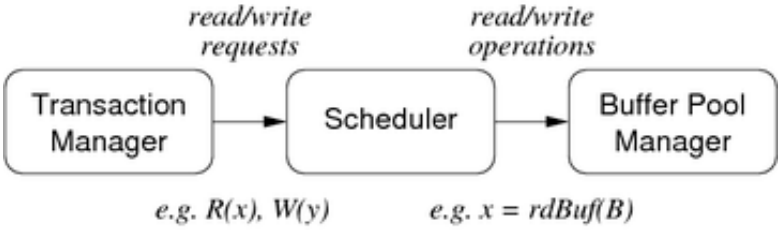
29/103

Simplest form of isolation: *serial* execution (T_1 ; T_2 ; T_3 ; ...)

Problem: serial execution yields poor throughput.

Concurrency control schemes (CCSs) aim for "safe" concurrency

Abstract view of DBMS concurrency mechanisms:



Serializability

30/103

Consider two schedules S_1 and S_2 produced by

- executing the same set of transactions $T_1..T_n$ concurrently
- but with a non-serial interleaving of *R/W* operations

S_1 and S_2 are *equivalent* if $StateAfter(S_1) = StateAfter(S_2)$

- i.e. final state yielded by S_1 is same as final state yielded by S_2

S is a *serializable schedule* (for a set of concurrent tx's $T_1 .. T_n$) if

- S is equivalent to some serial schedule S_s of $T_1 .. T_n$

Under these circumstances, consistency is guaranteed
(assuming no aborted transactions and no system failures)

... Serializability

31/103

Two formulations of serializability:

- *conflict serializibility*
  - i.e. conflicting R/W operations occur in the "right order"
  - check via precedence graph; look for absence of cycles
- *view serializibility*
  - i.e. read operations *see* the correct version of data
  - checked via VS conditions on likely equivalent schedules

View serializability is strictly weaker than conflict serializability.

Exercise 4: Serializability Checking

32/103

Is the following schedule view/conflict serializable?

```
T1:         W(B)  W(A)
T2:  R(B)                     W(A)
T3:                     R(A)        W(A)
```

Is the following schedule view/conflict serializable?

```
T1:         W(B)  W(A)
T2:  R(B)               W(A)
T3:                           R(A)  W(A)
```

Transaction Isolation Levels

33/103

SQL programmers' concurrency control mechanism ...

```
set transaction
    read only  -- so weaker isolation may be ok
    read write -- suggests stronger isolation needed
isolation level
    -- weakest isolation, maximum concurrency
    read uncommitted
    read committed
    repeatable read
```

```
serializable
-- strongest isolation, minimum concurrency
```

Applies to current tx only; affects how scheduler treats this tx.

... Transaction Isolation Levels

34/103

Implication of transaction isolation levels:

Isolation Level	Dirty Read	Nonrepeatable Read	Phantom Read
Read Uncommitted	Possible	Possible	Possible
Read Committed	Not Possible	Possible	Possible
Repeatable Read	Not Possible	Not Possible	Possible
Serializable	Not Possible	Not Possible	Not Possible

... Transaction Isolation Levels

35/103

For transaction isolation, PostgreSQL

- provides syntax for all four levels
- treats *read uncommitted* as *read committed*
- *repeatable read* behaves *like serializable*
- default level is *read committed*

Note: cannot implement *read uncommitted* because of MVCC

For more details, see PostgreSQL Documentation section 13.2

- extensive discussion of semantics of `UPDATE`, `INSERT`, `DELETE`

... Transaction Isolation Levels

36/103

A PostgreSQL tx consists of a sequence of SQL statements:

`BEGIN` S_1`;` S_2`; ...` S_n`;` `COMMIT;`

Isolation levels affect view of DB provided to each S_i:

- in *read committed* ...
  - each S_i sees snapshot of DB at start of S_i
- in *repeatable read* and *serializable* ...
  - each S_i sees snapshot of DB at start of tx
  - serializable checks for extra conditions

Transactions fail if the system detects violation of isolation level.

... Transaction Isolation Levels

Example of *repeatable read* vs *serializable*

- table R(class,value) containing (1,10) (1,20) (2,100) (2,200)
- T1: X = sum(value) where class=1; insert R(2,X); commit
- T2: X = sum(value) where class=2; insert R(1,X); commit
- with *repeatable read*, both transactions commit, giving
  - updated table: (1,10) (1,20) (2,100) (2,200) (1,300) (2,30)
- with *serial* transactions, only one transaction commits
  - T1;T2 gives (1,10) (1,20) (2,100) (2,200) (2,30) (1,330)
  - T2;T1 gives (1,10) (1,20) (2,100) (2,200) (1,300) (2,330)
- PG recognises that committing both gives serialization violation

Implementing Concurrency Control

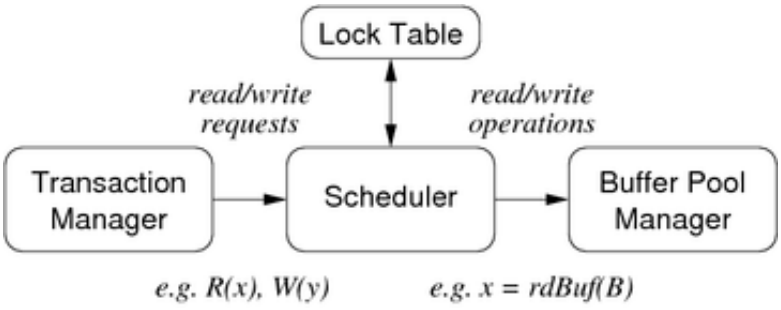
Concurrency Control

Approaches to concurrency control:

- *Lock-based*
  - Synchronise tx execution via locks on relevant part of DB.
- *Version-based* (multi-version concurrency control)
  - Allow multiple consistent versions of the data to exist.
    Each tx has access only to version existing at start of tx.
- *Validation-based* (optimistic concurrency control)
  - Execute all tx's; check for validity problems on commit.
- *Timestamp-based*
  - Organise tx execution via timestamps assigned to actions.

Lock-based Concurrency Control

Locks introduce additional mechanisms in DBMS:



The Lock Manager

- manages the locks requested by the scheduler

... Lock-based Concurrency Control

Lock table entries contain:

- object being locked (DB, table, tuple, field)
- type of lock: read/shared, write/exclusive

- FIFO queue of tx's requesting this lock
- count of tx's currently holding lock (max 1 for write locks)

Lock and unlock operations *must* be atomic.

Lock *upgrade*:

- if a tx holds a read lock, and it is the only tx holding that lock
- then the lock can be converted into a write lock

... Lock-based Concurrency Control

Synchronise access to shared data items via following rules:

- before reading *X*, get read (shared) lock on *X*
- before writing *X*, get write (exclusive) lock on *X*
- a tx attempting to get a read lock on *X* is blocked if another tx already has write lock on *X*
- a tx attempting to get an write lock on *X* is blocked if another tx has any kind of lock on *X*

These rules alone do not guarantee serializability.

... Lock-based Concurrency Control

Consider the following schedule, using locks:

```
T1(a): Lr(Y)     R(Y)           continued
T2(a):      Lr(X)    R(X) U(X)  continued

T1(b):      U(Y)          Lw(X) W(X) U(X)
T2(b): Lw(Y)....W(Y) U(Y)
```

(where L_r = read-lock, L_w = write-lock, U = unlock)

Locks correctly ensure controlled access to `X` and `Y`.

Despite this, the schedule is not serializable. (Ex: prove this)

Two-Phase Locking

To guarantee serializability, we require an additional constraint:

- in every tx, all *lock* requests precede all *unlock* requests

Each transaction is then structured as:

- *growing* phase where locks are acquired
- *action* phase where "real work" is done
- *shrinking* phase where locks are released

Clearly, this reduces potential concurrency ...

Problems with Locking

Appropriate locking can guarantee correctness.

However, it also introduces potential undesirable effects:

- *Deadlock*
  - No transactions can proceed; each waiting on lock held by another.
- *Starvation*
  - One transaction is permanently "frozen out" of access to data.
- *Reduced performance*
  - Locking introduces delays while waiting for locks to be released.

Deadlock

Deadlock occurs when two transactions are waiting for a lock on an item held by the other.

Example:

```
T1: L_w(A) R(A)              L_w(B) ......
T2:             L_w(B) R(B)         L_w(A) .....
```

How to deal with deadlock?

- prevent it happening in the first place
- let it happen, detect it, recover from it

... Deadlock

Handling deadlock involves forcing a transaction to "back off"

- select process to roll back
  - choose on basis of how far tx has progressed, # locks held, ...
- roll back the selected process
  - how far does this it need to be rolled back?
  - worst-case scenario: abort one transaction, then retry
- prevent starvation
  - need methods to ensure that same tx isn't always chosen

... Deadlock

Methods for managing deadlock

- *timeout* : set max time limit for each tx
- *waits-for graph* : records T_j waiting on lock held by T_k
  - *prevent* deadlock by checking for new cycle $\Rightarrow$ abort T_i
  - *detect* deadlock by periodic check for cycles $\Rightarrow$ abort T_i
- *timestamps* : use tx start times as basis for priority
  - scenario: T_j tries to get lock held by T_k ...
  - *wait-die*: if $T_j < T_k$, then T_j waits, else T_j rolls back
  - *wound-wait*: if $T_j < T_k$, then T_k rolls back, else T_j waits

---

... Deadlock

Properties of deadlock handling methods:

- both wait-die and wound-wait are fair
- wait-die tends to

  - roll back tx's that have done little work
  - but rolls back tx's more often
- wound-wait tends to
  - roll back tx's that may have done significant work
  - but rolls back tx's less often
- timestamps easier to implement than waits-for graph
- waits-for minimises roll backs because of deadlock

Exercise 5: Deadlock Handling

50/103

Consider the following schedule on four transactions:

```
T1:  R(A)        W(C)                                  W(D)
T2:       R(B)                          W(C)
T3:                    R(D)        W(B)
T4:                          R(E)              W(A)
```

Assume that: each `R` acquires a shared lock; each `W` uses an exclusive lock; two-phase locking is used.

Show how the wait-for graph for the locks evolves.

Show how any deadlocks might be resolved via this graph.

Transactions: the story so far

51/103

Transactions should obey ACID properties

Isolation can be compromised by uncontrolled concurrency

Serializable schedules avoid potential update anomalies

- less safe (more concurrent) isolation levels exist
- read uncommitted, read committed, repeatable read

Styles of concurrency control

- locking (two-phase, deadlock)
- optimistic concurrency control (try, then fix problems)
- multi-version concurrency control (less locking needed)

Optimistic Concurrency Control

52/103

Locking is a pessimistic approach to concurrency control:

- limit concurrency to ensure that conflicts don't occur

Costs: lock management, deadlock handling, contention.

In scenarios where there are far more reads than writes ...

- don't lock (allow arbitrary interleaving of operations)
- check just before commit that no conflicts occurred
- if problems, roll back conflicting transactions

Optimistic concurrency control (OCC) is a strategy to realise this.

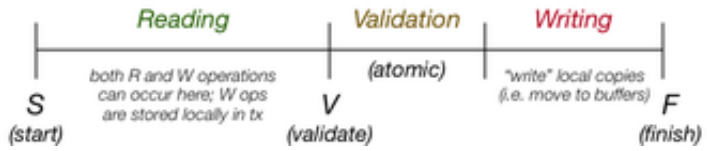
53/103

... Optimistic Concurrency Control

Under OCC, transactions have three distinct phases:

- *Reading*: read from database, modify local copies of data
- *Validation*: check for conflicts in updates
- *Writing*: commit local copies of data to database

Timestamps are recorded at points *S, V, F* :



... Optimistic Concurrency Control

Data structures needed for validation:

- *S* ... set of txs that are reading data and computing results
- *V* ... set of txs that have reached validation (not yet committed)
- *F* ... set of txs that have finished (committed data to storage)
- for each T_i, timestamps for when it reached *S*, *V*, *F*
- $RS(T_i)$ set of all data items read by T_i
- $WS(T_i)$ set of all data items to be written by T_i

Use the *V* timestamps as ordering for transactions

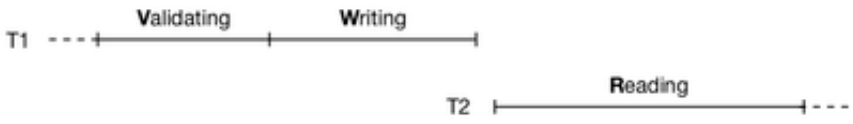
- assume serial tx order based on ordering of $V(T_i)$'s

... Optimistic Concurrency Control

Two-transaction example:

- allow transactions T_1 and T_2 to run without any locking
- check that objects used by T_2 are not being changed by T_1
- if they are, we need to roll back T_2 and retry

Case 0: serial execution ... no problem

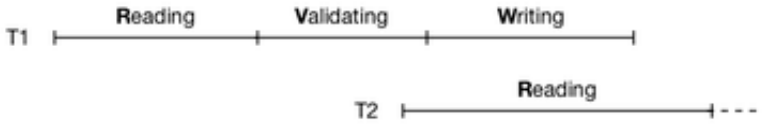


... Optimistic Concurrency Control

Case 1: reading overlaps validation/writing

- T_2 starts while T_1 is validating/writing
- if some *X* being read by T_2 is in $WS(T_1)$

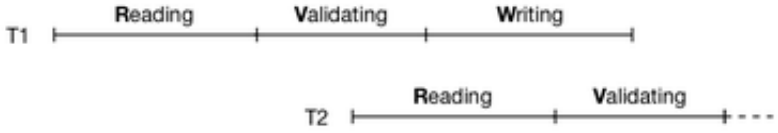
- then T_2 may not have read the updated version of X
- so, T_2 must start again



... Optimistic Concurrency Control

Case 2: reading/validation overlaps validation/writing

- T_2 starts validating while T_1 is validating/writing
- if some X being written by T_2 is in $WS(T_1)$
- then T_2 may end up overwriting T_1's update
- so, T_2 must start again



... Optimistic Concurrency Control

Validation check for transaction T

- for all transactions $T_i \neq T$
  - if $T \in S$ & $T_i \in F$, then ok
  - if $T \notin V$ & $V(T_i) < S(T) < F(T_i)$,
    then check $WS(T_i) \cap RS(T)$ is empty
  - if $T \in V$ & $V(T_i) < V(T) < F(T_i)$,
    then check $WS(T_i) \cap WS(T)$ is empty

If this check fails for any T_i, then T is rolled back.

... Optimistic Concurrency Control

OCC prevents: T reading dirty data, T overwriting T_i's changes

Problems with OCC:

- increased roll backs**
- tendency to roll back "complete" tx's
- cost to maintain S,V,F sets

** "Roll back" is relatively cheap

- changes to data are purely local before Writing phase
- no requirement for logging info or undo/redo (see later)

Multi-version Concurrency Control

Multi-version concurrency control (MVCC) aims to

- retain benefits of locking, while getting more concurrency
- by providing multiple (consistent) versions of data items

Achieves this by

- readers access an "appropriate" version of each data item
- writers make new versions of the data items they modify

Main difference between MVCC and standard locking:

- read locks do not conflict with write locks ⇒
- reading never blocks writing, writing never blocks reading

... Multi-version Concurrency Control

WTS = timestamp of tx that wrote this data item

Chained tuple versions: $tup_{oldest} \rightarrow tup_{older} \rightarrow tup_{newest}$

When a reader T_i is accessing the database

- ignore any data item D created after T_i started
  - checked by: WTS(D) > TS(T_i)
- use only newest version V accessible to T_i
  - determined by: max(WTS(V)) < TS(T_i)

... Multi-version Concurrency Control

When a writer T_i attempts to change a data item

- find newest version V satisfying WTS(V) < TS(T_i)
- if no later versions exist, create new version of data item
- if there are later versions, then abort T_i

Some MVCC versions also maintain RTS (TS of last reader)

- don't allow T_i to write D if RTS(D) > TS(T_i)

... Multi-version Concurrency Control

Advantage of MVCC

- locking needed for serializability considerably reduced

Disadvantages of MVCC

- visibility-check overhead (on every tuple read/write)
- reading an item *V* causes an update of *RTS(V)*
- storage overhead for extra versions of data items
- overhead in removing out-of-date versions of data items

Despite apparent disadvantages, MVCC is very effective.

... Multi-version Concurrency Control

Removing old versions:

- V_j and V_k are versions of same item
- $WTS(V_j)$ and $WTS(V_k)$ precede $TS(T_i)$ for all T_i
- remove version with smaller $WTS(V_x)$ value

When to make this check?

- every time a new version of a data item is added?
- periodically, with fast access to blocks of data

PostgreSQL uses the latter (*vacuum*).

Concurrency Control in PostgreSQL

PostgreSQL uses two styles of concurrency control:

- multi-version concurrency control (MVCC)
  (used in implementing SQL DML statements (e.g. `select`))
- two-phase locking (2PL)
  (used in implementing SQL DDL statements (e.g. `create table`))

From the SQL (PLpgSQL) level:

- can let the lock/MVCC system handle concurrency
- can handle it explicitly via `LOCK` statements

... Concurrency Control in PostgreSQL

PostgreSQL provides *read committed* and *serializable* isolation levels.

Using the serializable isolation level, a `select`:

- sees only data committed before the transaction began
- never sees changes made by concurrent transactions

Using the serializable isolation level, an update fails:

- if it tries to modify an "active" data item
  (active = affected by some other tx, either committed or uncommitted)

The transaction containing the update must then rollback and re-start.

... Concurrency Control in PostgreSQL

Implementing MVCC in PostgreSQL requires:

- a log file to maintain current status of each T_i
- in every tuple:
  - `xmin` ID of the tx that created the tuple

  - `xmax` ID of the tx that replaced/deleted the tuple (if any)
  - `xnew` link to newer versions of tuple (if any)
- for each transaction T_i :
  - a transaction ID (timestamp)
  - SnapshotData: list of active tx's when T_i started

... Concurrency Control in PostgreSQL

Rules for a tuple to be visible to T_i :

- the `xmin` (creation transaction) value must
  - be committed in the log file
  - have started before T_i's start time
  - not be active at T_i's start time
- the `xmax` (delete/replace transaction) value must
  - be blank or refer to an aborted tx, or
  - have started after T_i's start time, or
  - have been active at SnapshotData time

For details, see: **`utils/time/tqual.c`**

... Concurrency Control in PostgreSQL

Tx's always see a consistent version of the database.

But may not see the "current" version of the database.

E.g. T_1 does select, then concurrent T_2 deletes some of T_1's selected tuples

This is OK unless tx's communicate outside the database system.

E.g. T_1 counts tuples; T_2 deletes then counts; then counts are compared

Use locks if application needs every tx to see same current version

- `LOCK TABLE` locks an entire table
- `SELECT FOR UPDATE` locks only the selected rows

Exercise 6: Locking in PostgreSQL

How could we solve this problem via locking?

```
create or replace function
    allocSeat(paxID int, fltID int, seat text)
    returns boolean
as $$
declare
    pid int;
begin
    select paxID into pid from SeatingAlloc
    where  flightID = fltID and seatNum = seat;
    if (pid is not null) then
        return false;  -- someone else already has seat
    else
        update SeatingAlloc set pax = paxID
        where  flightID = fltID and seatNum = seat;
        commit;
```

```
        return true;
    end if;
end;
$$ langauge plpgsql;
```

Implementing Atomicity/Durability

Atomicity/Durability

72/103

Reminder:

Transactions are *atomic*

- if a tx commits, all of its changes persist in DB
- if a tx aborts, none of its changes occur in DB

Transaction effects are *durable*

- if a tx commits, its effects persist
  (even in the event of subsequent (catastrophic) system failures)

Implementation of atomicity/durability is intertwined.

Durability

73/103

What kinds of "system failures" do we need to deal with?

- single-bit inversion during transfer mem-to-disk
- decay of storage medium on disk (some data changed)
- failure of entire disk device (data no longer accessible)
- failure of DBMS processes (e.g. `postgres` crashes)
- operating system crash; power failure to computer room
- complete destruction of computer system running DBMS

The last requires off-site *backup*; all others should be locally recoverable.

... Durability

74/103

Consider following scenario:

Desired behaviour after system restart:

- all effects of T1, T2 persist
- as if T3, T4 were aborted (no effects remain)

... Durability

Durabilty begins with a *stable disk storage subsystem*

- i.e. `putPage()` and `getPage()` always work as expected

We can prevent/minimise loss/corruption of data due to:

- mem/disk transfer corruption ⇒ parity checking
- sector failure ⇒ mark "bad" blocks
- disk failure ⇒ RAID (levels 4,5,6)
- destruction of computer system ⇒ off-site backups

Dealing with Transactions

The remaining "failure modes" that we need to consider:

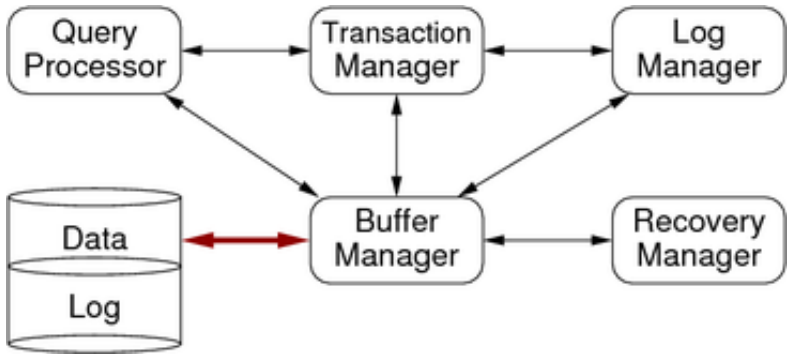
- failure of DBMS processes or operating system
- failure of transactions (`ABORT`)

Standard technique for managing these:

- keep a *log* of changes made to database
- use this log to restore state in case of failures

Architecture for Atomicity/Durability

How does a DBMS provide for atomicity/durability?



Execution of Transactions

Transactions deal with three address spaces:

- stored data on the disk (representing global DB state)
- data in memory buffers (where held for sharing by tx's)
- data in their own local variables (where manipulated)

Each of these may hold a different "version" of a DB object.

PostgreSQL processes make heavy use of shared buffer pool

⇒ transactions do not deal with much local data.

... Execution of Transactions

79/103

Operations available for data transfer:

- `INPUT(X)` ... read page containing `X` into a buffer
- `READ(X,v)` ... copy value of `X` from buffer to local var `v`
- `WRITE(X,v)` ... copy value of local var `v` to `X` in buffer
- `OUTPUT(X)` ... write buffer containing `X` to disk

`READ/WRITE` are issued by transaction.

`INPUT/OUTPUT` are issued by buffer manager (and log manager).

`INPUT/OUTPUT` correspond to `getPage()`/`putPage()` mentioned above

... Execution of Transactions

80/103

Example of transaction execution:

```
-- implements A = A*2; B = B+1;
BEGIN
READ(A,v); v = v*2; WRITE(A,v);
READ(B,v); v = v+1; WRITE(B,v);
COMMIT
```

`READ` accesses the buffer manager and may cause `INPUT`.

`COMMIT` needs to ensure that buffer contents go to disk.

... Execution of Transactions

81/103

States as the transaction executes:

```
t   Action        v  Buf(A)  Buf(B)  Disk(A)  Disk(B)
-------------------------------------------------------
(0) BEGIN         .      .       .        8        5
(1) READ(A,v)     8      8       .        8        5
(2) v = v*2      16      8       .        8        5
(3) WRITE(A,v)   16     16       .        8        5
(4) READ(B,v)     5     16       5        8        5
(5) v = v+1       6     16       5        8        5
(6) WRITE(B,v)    6     16       6        8        5
(7) OUTPUT(A)     6     16       6       16        5
(8) OUTPUT(B)     6     16       6       16        6
```

After tx completes, we must have either
Disk(A)=8, Disk(B)=5 or Disk(A)=16, Disk(B)=6

If system crashes before (8), may need to undo disk changes.
If system crashes after (8), may need to redo disk changes.

Transactions and Buffer Pool

82/103

Two issues arise w.r.t. buffers:

- *forcing* ... `OUTPUT` buffer on each `WRITE`
  - ensures durability; disk always consistent with buffer pool
  - poor performance; defeats purpose of having buffer pool
- *stealing* ... replace buffers of uncommitted tx's
  - if we don't, poor throughput (tx's blocked on buffers)
  - if we do, seems to cause atomicity problems?

Ideally, we want stealing and not forcing.

... Transactions and Buffer Pool

83/103

Handling *stealing*:

- transaction T loads page P and makes changes
- T_2 needs a buffer, and P is the "victim"
- P is output to disk (it's dirty) and replaced
- if T aborts, some of its changes are already "committed"
- must log values changed by T in P at "steal-time"
- use these to UNDO changes in case of failure of T

... Transactions and Buffer Pool

84/103

Handling *no forcing*:

- transaction T makes changes & commits, then system crashes
- but what if modified page P has not yet been output?
- must log values changed by T in P as soon as they change
- use these to support REDO to restore changes

Above scenario may be a problem, even if we are forcing

- e.g. system crashes immediately after requesting a `WRITE()`

Logging

85/103

Three "styles" of logging

- *undo* ... removes changes by any uncommitted tx's
- *redo* ... repeats changes by any committed tx's
- *undo/redo* ... combines aspects of both

All approaches require:

- a sequential file of log records
- each log record describes a change to a data item
- log records are written first
- actual changes to data are written later

Known as *write-ahead logging* (PostgreSQL uses WAL)

Undo Logging

86/103

Simple form of logging which ensures atomicity.

Log file consists of a *sequence* of small records:

- <START T> ... transaction T begins
- <COMMIT T> ... transaction T completes successfully
- <ABORT T> ... transaction T fails (no changes)
- <T,X,v> ... transaction T changed value of X from v

Notes:

- we refer to <T,X,v> generically as <UPDATE> log records
- update log entry created for each WRITE (not OUTPUT)
- update log entry contains *old* value (new value is not recorded)

... Undo Logging

87/103

Data must be written to disk in the following order:

1. <START> transaction log record
2. <UPDATE> log records indicating changes
3. the changed data elements themselves
4. <COMMIT> log record

Note: sufficient to have <T,X,v> output before X, for each X

... Undo Logging

88/103

For the example transaction, we would get:

```
t    Action        v  B(A)  B(B)  D(A)  D(B)  Log
-----------------------------------------------------
(0)  BEGIN         .    .     .     8     5   <START T>
(1)  READ(A,v)     8    8     .     8     5
(2)  v = v*2      16    8     .     8     5
(3)  WRITE(A,v)   16   16     .     8     5   <T,A,8>
(4)  READ(B,v)     5   16     5     8     5
(5)  v = v+1       6   16     5     8     5
(6)  WRITE(B,v)    6   16     6     8     5   <T,B,5>
(7)  FlushLog
(8)  StartCommit
(9)  OUTPUT(A)     6   16     6    16     5
(10) OUTPUT(B)     6   16     6    16     6
(11) EndCommit                                <COMMIT T>
(12) FlushLog
```

Note that T is not regarded as committed until (12) completes.

... Undo Logging

89/103

Simplified view of recovery using UNDO logging:

- scan *backwards* through log
  - if <COMMIT T>, mark T as committed
  - if <T,X,v> and T not committed, set X to v on disk
  - if <START T> and T not committed, put <ABORT T> in log

Assumes we scan entire log; use checkpoints to limit scan.

... Undo Logging

Algorithmic view of recovery using UNDO logging:

```
committedTrans = abortedTrans = startedTrans = {}
for each log record from most recent to oldest {
    switch (log record) {
    <COMMIT T> : add T to committedTrans
    <ABORT T>  : add T to abortedTrans
    <START T>  : add T to startedTrans
    <T,X,v>    : if (T in committedTrans)
                     // don't undo committed changes
                 else  // roll-back changes
                     { WRITE(X,v); OUTPUT(X) }
}   }
for each T in startedTrans {
    if (T in committedTrans) ignore
    else if (T in abortedTrans) ignore
    else write <ABORT T> to log
}
flush log
```

Checkpointing

Simple view of recovery implies reading entire log file.

Since log file grows without bound, this is infeasible.

Eventually we can delete "old" section of log.

- i.e. where *all* prior transactions have committed

This point is called a *checkpoint*.

- all of log prior to checkpoint can be ignored for recovery

... Checkpointing

Problem: many concurrent/overlapping transactions.

How to know that all have finished?

1. periodically, write log record <CHKPT (T1,..,Tk)>
   (contains references to all active transactions ⇒ active tx table)
2. continue normal processing (e.g. new tx's can start)
3. when all of T1,..,Tk have completed,
   write log record <ENDCHKPT> and flush log

Note: tx manager maintains chkpt and active tx information

... Checkpointing

Recovery: scan backwards through log file processing as before.

Determining where to stop depends on ...

- whether we meet <ENDCHKPT> or <CHKPT...> first

If we encounter <ENDCHKPT> first:

- we know that all incomplete tx's come after prev <CHKPT...>
- thus, can stop backward scan when we reach <CHKPT...>

If we encounter <CHKPT (T1,..,Tk)> first:

- crash occurred *during* the checkpoint period
- any of T1,..,Tk that committed before crash are ok
- for uncommitted tx's, need to continue backward scan

Redo Logging

94/103

Problem with UNDO logging:

- all changed data must be output to disk before committing
- conflicts with optimal use of the buffer pool

Alternative approach is *redo* logging:

- allow changes to remain only in buffers after commit
- write records to indicate what changes are "pending"
- after a crash, can apply changes during recovery

... Redo Logging

95/103

Requirement for redo logging: *write-ahead rule*.

Data must be written to disk as follows:

1. start transaction log record
2. update log records indicating changes
3. then commit log record (OUTPUT)
4. then OUTPUT changed data elements themselves

Note that update log records now contain <T,X,v'>,
where v' is the *new* value for X.

... Redo Logging

96/103

For the example transaction, we would get:

```
t    Action        v  B(A)  B(B)  D(A)  D(B)  Log
-----------------------------------------------------
(0)  BEGIN         .    .     .     8     5   <START T>
(1)  READ(A,v)     8    8     .     8     5
(2)  v = v*2      16    8     .     8     5
(3)  WRITE(A,v)   16   16     .     8     5   <T,A,16>
(4)  READ(B,v)     5   16     5     8     5
(5)  v = v+1       6   16     5     8     5
(6)  WRITE(B,v)    6   16     6     8     5   <T,B,6>
(7)  COMMIT                                  <COMMIT T>
(8)  FlushLog
(9)  OUTPUT(A)     6   16     6    16     5
(10) OUTPUT(B)     6   16     6    16     6
```

Note that T is regarded as committed as soon as (8) completes.

... Redo Logging

Simplified view of recovery using REDO logging:

- identify all committed tx's (backwards scan)
- scan *forwards* through log
  - if <T,X,v> and T is committed, set X to v on disk
  - if <START T> and T not committed, put <ABORT T> in log

Assumes we scan entire log; use checkpoints to limit scan.

Undo/Redo Logging

UNDO logging and REDO logging are incompatible in

- order of outputting <COMMIT T> and changed data
- how data in buffers is handled during checkpoints

Undo/Redo logging combines aspects of both

- requires new kind of update log record
  <T,X,v,v'> gives both old and new values for X
- removes incompatibilities between output orders

As for previous cases, requires write-ahead of log records.

Undo/redo loging is common in practice; Aries algorithm.

... Undo/Redo Logging

For the example transaction, we might get:

```
t    Action        v  B(A)  B(B)  D(A)  D(B)  Log
-----------------------------------------------------
(0)  BEGIN         .    .     .     8     5   <START T>
(1)  READ(A,v)     8    8     .     8     5
(2)  v = v*2      16    8     .     8     5
(3)  WRITE(A,v)   16   16     .     8     5   <T,A,8,16>
(4)  READ(B,v)     5   16     5     8     5
(5)  v = v+1       6   16     5     8     5
(6)  WRITE(B,v)    6   16     6     8     5   <T,B,5,6>
(7)  FlushLog
(8)  StartCommit
(9)  OUTPUT(A)     6   16     6    16     5
(10)                                        <COMMIT T>
(11) OUTPUT(B)     6   16     6    16     6
```

Note that T is regarded as committed as soon as (10) completes.

... Undo/Redo Logging

Simplified view of recovery using UNDO/REDO logging:

- scan log to determine committed/uncommitted txs
- for each uncommitted tx T add <ABORT T> to log

- scan *backwards* through log
  - if <T,X,v,w> and T is not committed, set X to v on disk
- scan *forwards* through log
  - if <T,X,v,w> and T is committed, set X to w on disk

... Undo/Redo Logging

The above description simplifies details of undo/redo logging.

Aries is a complete algorithm for undo/redo logging.

Differences to what we have described:

- log records contain a sequence numnber (LSN)
- LSNs used in tx and buffer managers, and stored in data pages
- additional log record to mark <END> (of commit or abort)
- <CHKPT> contains only a timestamp
- <ENDCHKPT..> contains tx and dirty page info

Recovery in PostgreSQL

PostgreSQL uses write-ahead undo/redo style logging.

It also uses multi-version concurrency control, which

- tags each record with a tx and update timestamp

MVCC simplifies some aspects of undo/redo, e.g.

- some info required by logging is already held in each tuple
- no need to undo effects of aborted tx's; use old version

... Recovery in PostgreSQL

Transaction/logging code is distributed throughout backend.

Core transaction code is in **src/backend/access/transam**.

Transaction/logging data is written to files in **PGDATA/pg_xlog**

- a number of very large files containing log records
- old files are removed once all txs noted there are completed
- new files added when existing files reach their capacity (16MB)
- number of tx log files varies depending on tx activity

Produced: 1 Aug 2019