Raft & Spanner

IPADS, Shanghai Jiao Tong University

https://www.sjtu.edu.cn

Credits: John Ousterhout

Review: Replicated State Machines

A general approach to making consistent replicas of a server:

- Start with the **same initial state** on each server
- Provide each replica with the **same input** operations, in **same order**
- Ensure all operations are **deterministic**
  - E.g., no randomness, no reading of current time, etc.

These rules ensure each server will end up in the same final state

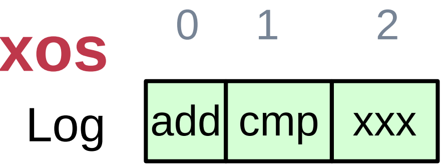
Review: Single-decree Paxos vs. Multi-Paxos

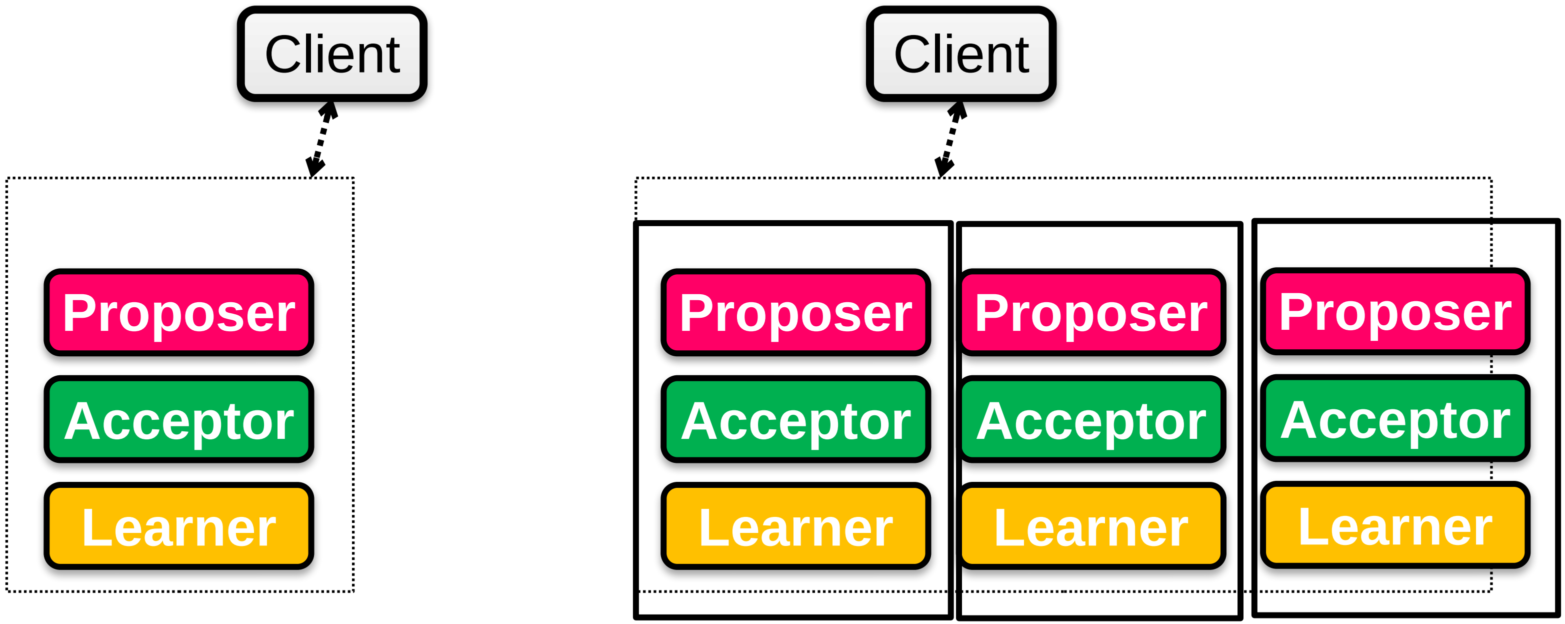
(Single-decree) Paxos allows us to ensure a consistent value over acceptors

- Value can be anything, e.g., an integer or a command

Multi-paxos adopts multiple single-decree paxos instances to realize logs in replicated state machine

Review: Single-decree Paxos vs. Multi-Paxos





Single-decree Paxos

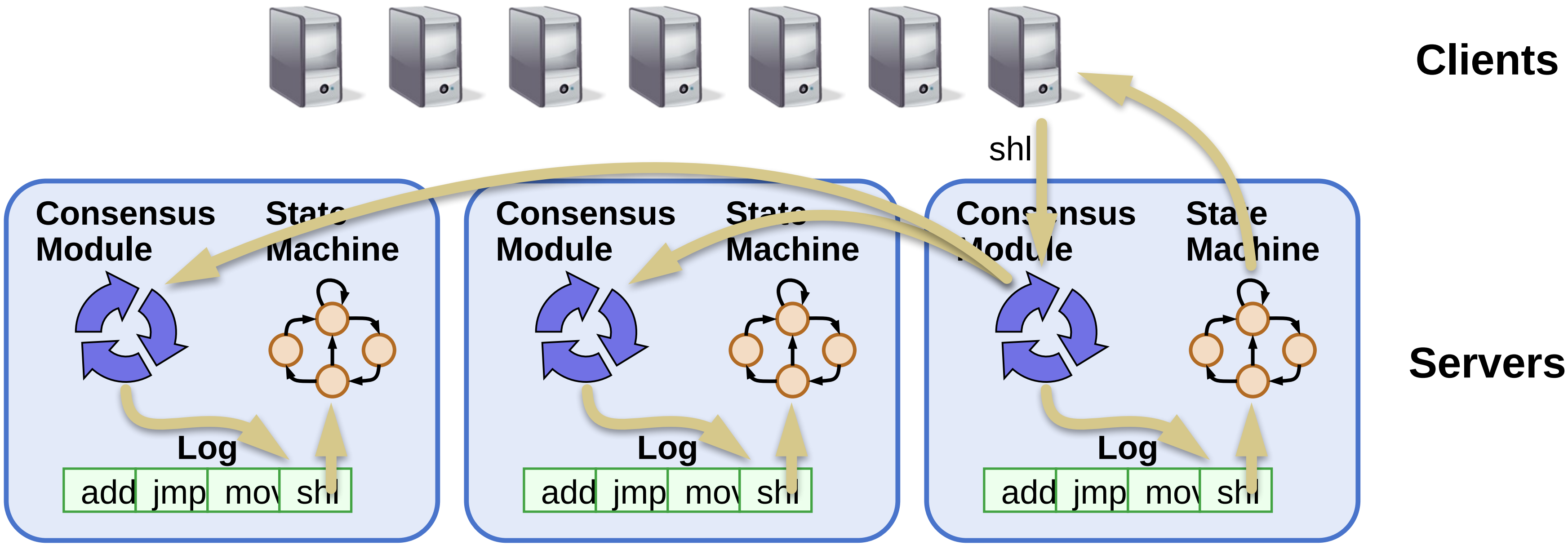
Single-decree instance for 0

Single-decree instance for 1

Single-decree instance for 2

Multi-Paxos

Review: raft's replicated state machine based on log



Replicated log => replicated state machine

- All servers execute same (deterministic) commands in same order

Consensus module ensures proper logs are the same!

Raft’s high-level approach: problem decomposition

1. **Leader election**

   Select one server as the leader

   Detect crashes, choose new leader

2. **Log replication (normal operation)**

   Leader accepts commands from clients, append to its log

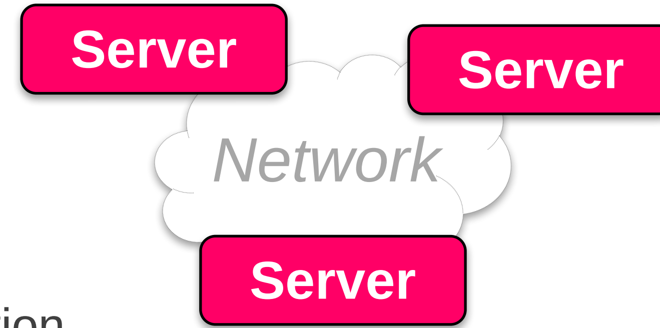
   Leader replicates its log to other servers (overwrites inconsistencies)

3. **Safety**

   Keep logs consistent

   Only servers with up-to-date logs can become the leader

Review: Raft server states



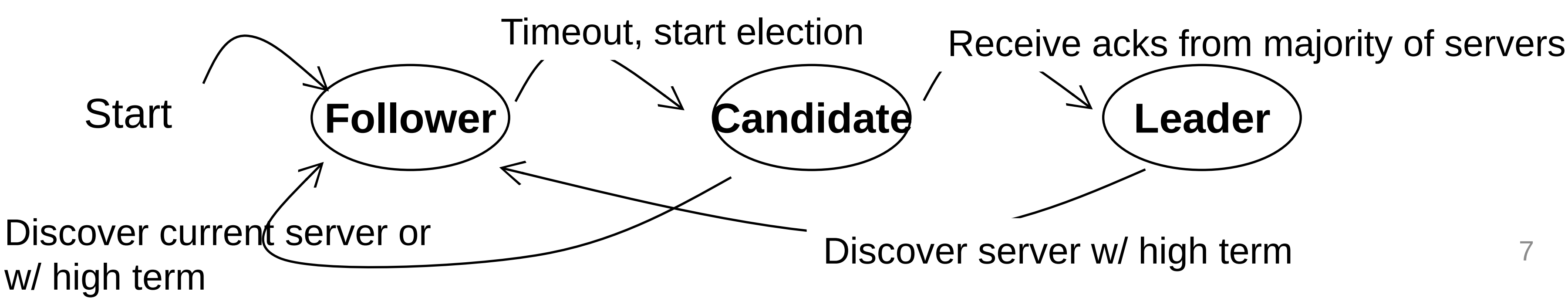
Servers communicates w/ RPCs

At any time, each server is either:

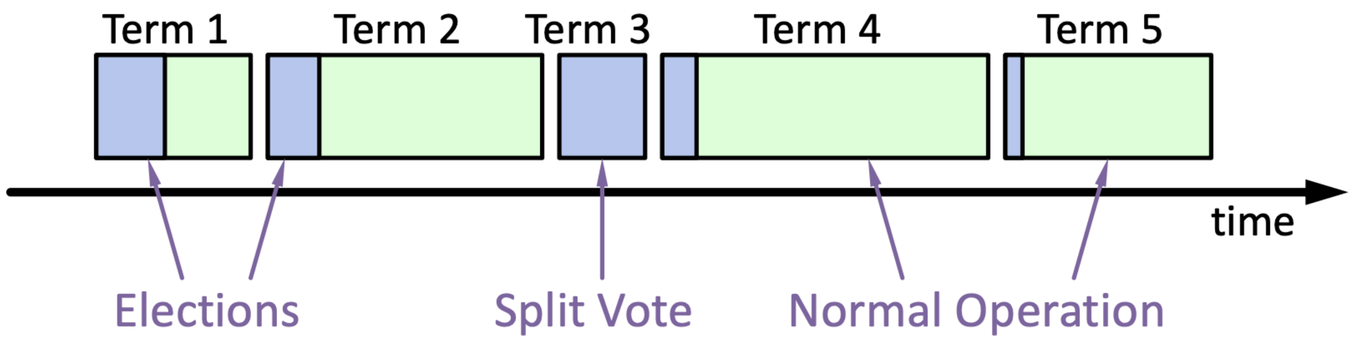
- **Leader**: handles all client interactions, log replication
  - Invariant: At most 1 viable leader at a time
- **Follower**: passive (only responds to incoming RPCs)
- **Candidate**: used to elect a new leader

Normal workloads

- 1 server is the leader, others are the followers



Review of Raft basics: terms for one leader



Raft divides time into terms (with arbitrary length):

- Each term starts with an election
- Ends with one leader or no leader
- At most one leader per term

Each leader is uniquely associated with a term

- Key role: identify obsolete information

Review: Basic request vote RPC so far

Invoked by candidates to gather votes.

Arguments:

candidateId candidate requesting vote

term candidate's term

Results:

term currentTerm, for candidate to update itself

voteGranted true means candidate received vote

Implementation:

1. If term > currentTerm, currentTerm ← term
   (step down if leader or candidate)
2. If term == currentTerm, votedFor is null or candidateId, grant vote and reset election timeout

Raft's high-level approach: problem decomposition

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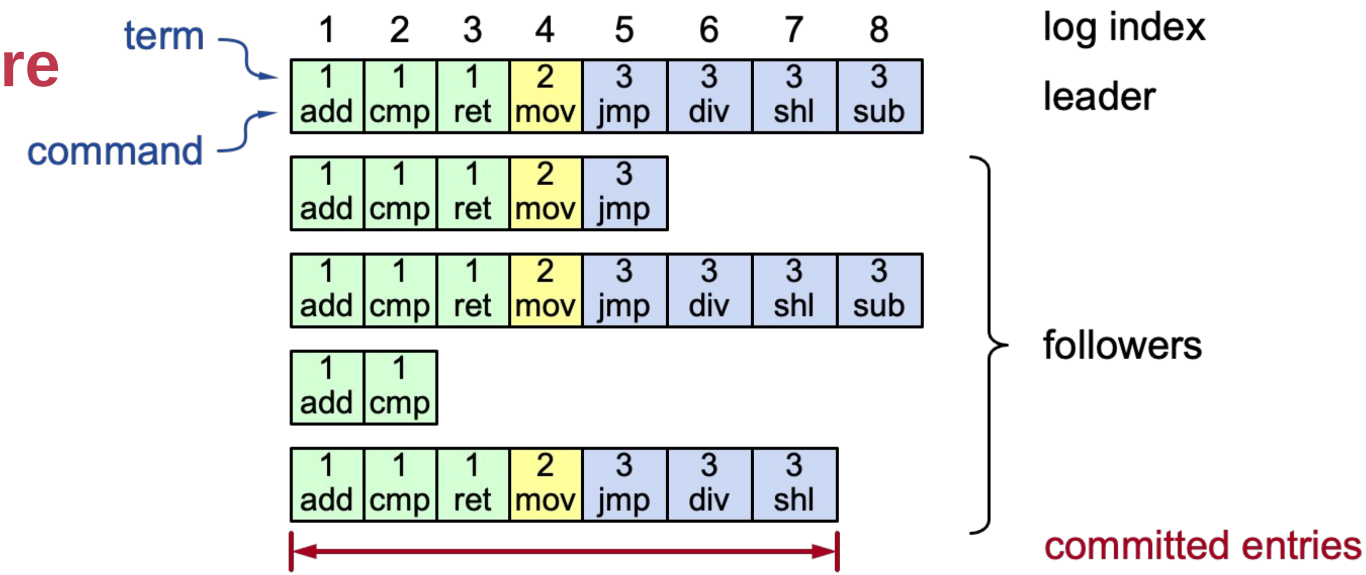
Log structure



Log entry = index, term, command

– Stored on the disk to tolerate failures

A log is committed if it can be safely applied to the state machine

– i.e., eventually stored on all the servers with the same value

Not all **entries are** **committed** (will talk about later)

Persistent state of each server + log

currentTerm

Latest term server has seen (initialized to 0 on first boot)

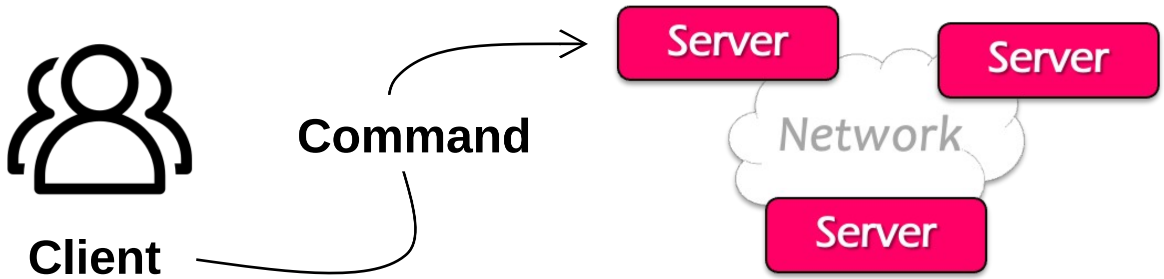
votedFor

Candidate Id that received vote in current term (or null if none)

Log[]

Log entries

Normal operations to append the log



① Send command to the leader

② Leader appends command to its log

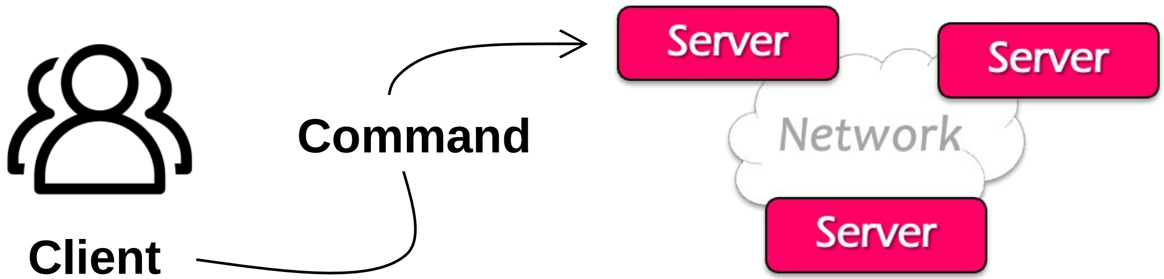
③ Leader sends **`AppendEntries`** RPCs to followers

④ Once a new entry (of log) **committed**:

Leader passes command to its state machine, returns results to client

Notifies followers of committed entries, Follower pass committed commands to their state machines (How to determine an entry is committed will be talked later)

Normal operations to update the log



Crashed/slow followers?

– Leader retries RPCs until they succeed (at least once)

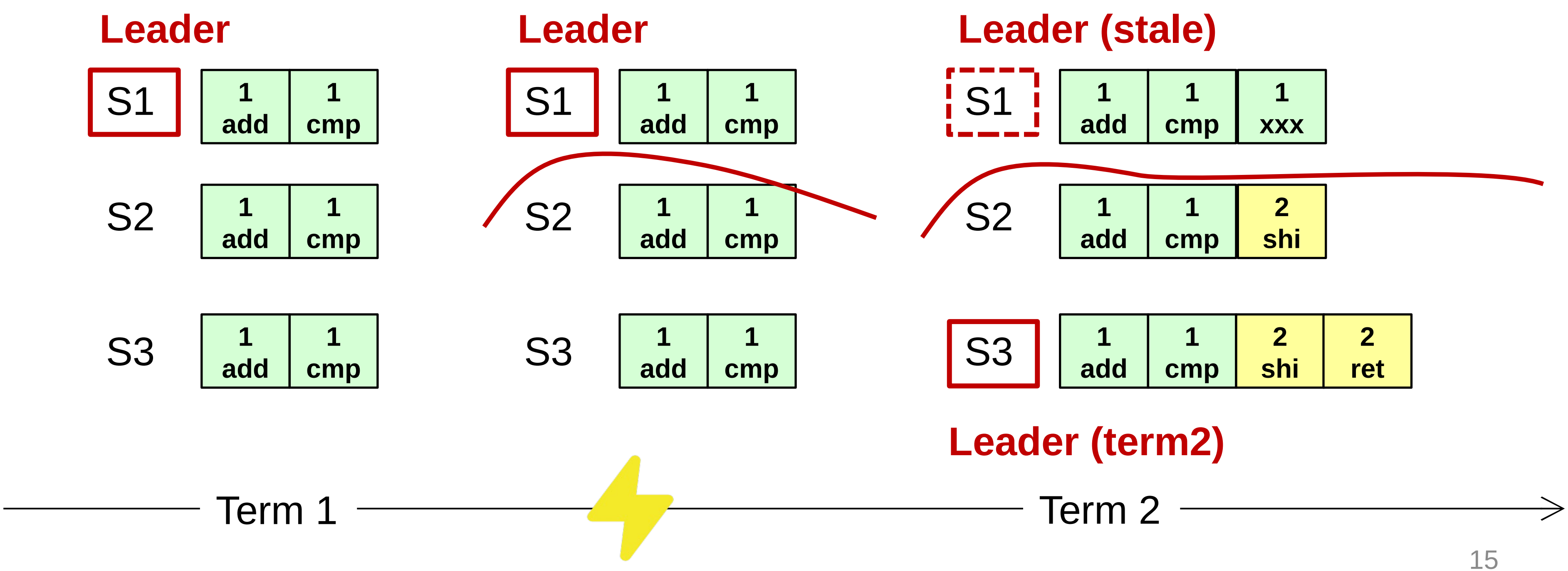
Performance is optimal in the common case

– One successful RPC to any majority of servers

Challenge: crash can cause log to inconsistencies

Case: an old leader is unaware of the new leader due to network partition

- Yet, it can still get commands from the clients



Challenge: crash can cause log to inconsistencies

	1	2	3	4	5	6	7	8	9	10	log index
s_1	1 x←3	1 q←8	1 j←2	2 x←q	2 z←5	3 y←1	3 y←3	3 q←j			leader for term 4
s_2	1 x←3	1 q←8	1 j←2	2 x←q	2 z←5	3 y←1	3 y←3				
s_3	1 x←3	1 q←8	1 j←2	2 x←q	2 z←5	3 y←1	3 y←3	3 q←j	3 x←4		followers
s_4	1 x←3	1 q←8									
s_5	1 x←3	1 q←8	1 j←2	2 x←q	2 z←5	2 y←3	2 q←j	2 x←8	2 x←4		

Raft minimizes special code for repairing inconsistencies

– Leaders assume its log is correct

– Normal operation will repair all inconsistencies

Consistency of the log

Question: how to achieve this property?

High level of coherency between logs maintained by the raft:

– If log entries on different servers have the same index & term
  - They store the same command
  - The logs are identical in all preceding entries

1	2	3	4	5	6
1 add	1 cmp	1 ret	2 mov	3 jmp	3 div

1	2	3	4	5	6
1 add	1 cmp	1 ret	2 mov	3 jmp	4 sub

If a given entry is committed, all preceding entries are also committed

– Note that not all log entries are committed

`AppendEntries` consistency checks

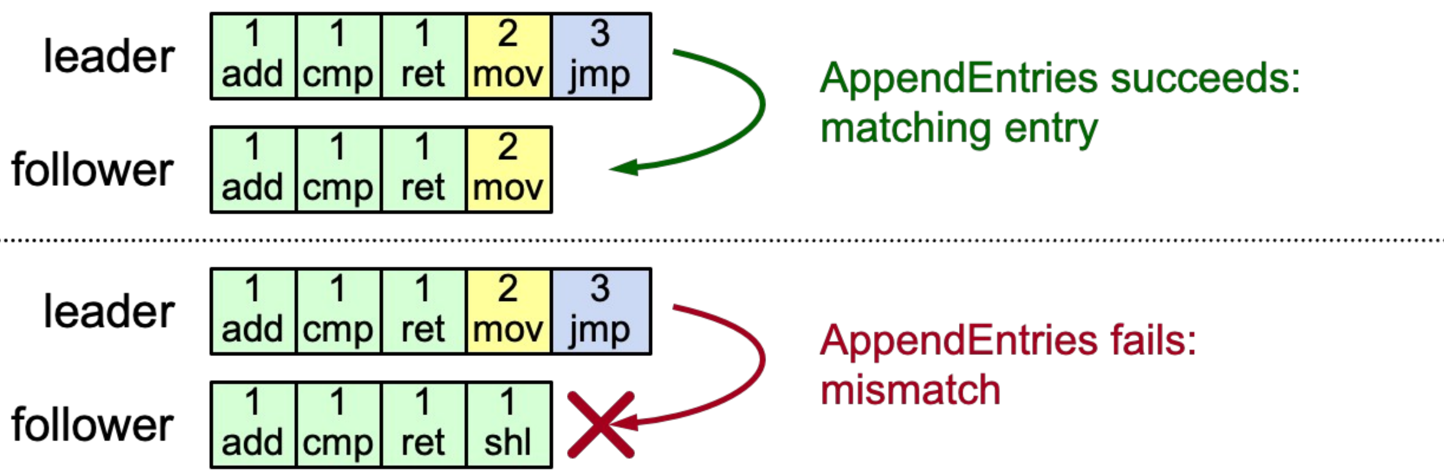
Implements an induction step, ensures coherency

Each RPC argument contains

- Append index, term, **term of entry preceding new ones**

Follower checks whether it has the matching entry

- Otherwise, it rejects the request



Inconsistent entries causes: leader changes or missing entries

At beginning of new leader's term:

- Old leader may have left entries partially replicated

Missing entries: a replica may be fall behind

No special steps by new leader: just start normal operation, w/ different position

- Leader's log is "the truth"
- Will eventually make follower's logs identical to leader's (overwrite the divergent log entries with the leader's ones)

The rejected/missing entries will be overwritten by the leader's log entry

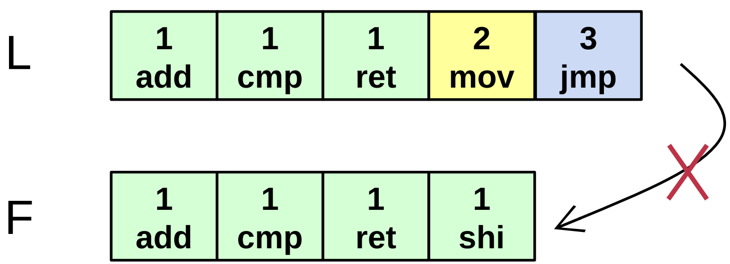
Handle rejections

Implements an induction step, ensures coherency

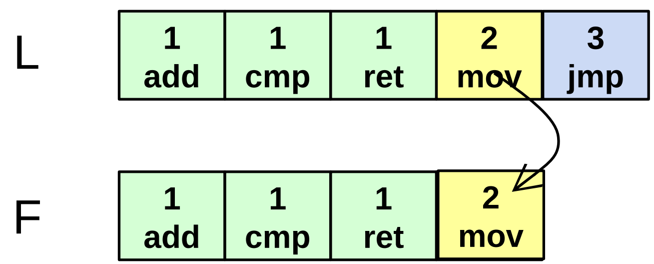
In this example, the leader will

- Overwrite index 4 with [2 mov] with AppendEntries
- Then append entry [3 jmp] to the end of the follower

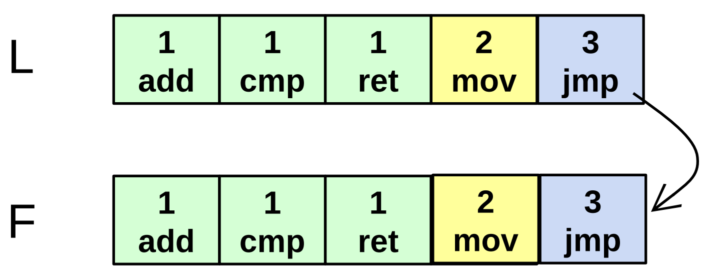
Fail to append case



Overwrite mismatched entry



Append the new entry



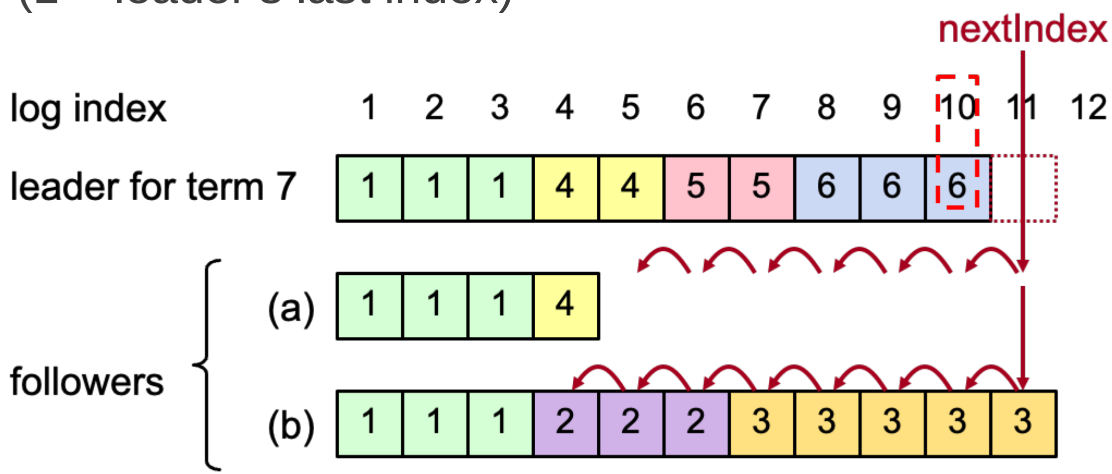
More on Repairing Follower Logs

New leader must make follower logs consistent with its own

- Delete extraneous entries
- Fill in missing entries

Leader keeps `nextIndex` for each follower:

- Index of next log entry to send to that follower
- Initialized to (1 + leader's last index)



AppendEntries RPC (Simplified)

Invoked by leader to replicate log entries and discover inconsistencies

Arguments:

term leader's term

prevLogIndex index of log entry immediately preceding new ones

prevLogTerm term of prevLogIndex entry

entries[] log entries to store (empty for heartbeat)

Results:

success true if follower contained entry matching prevLogIndex and prevLogTerm

AppendEntries RPC (Simplified)

Implementation:

1. Return false if term < currentTerm
2. If term > currentTerm, currentTerm ← term
3. If candidate or leader, step down
4. Reset election timeout
5. Return failure if log doesn't contain an entry at prevLogIndex whose term matches prevLogTerm
6. If existing entries conflict with new entries, delete all existing entries starting with first conflicting entry
7. Append any new entries not already in the log

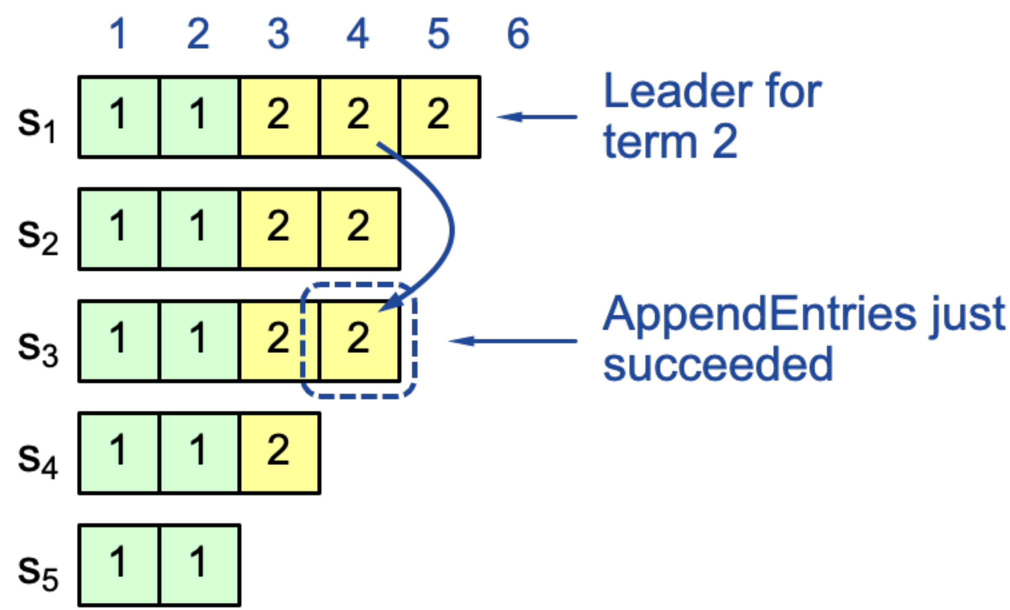
Overwrite makes a subtle issue:
When can we commit a log entry? (Since an appended log entry may be overwritten)

Intuitive goal: replicating on a majority

If the log is replicated on a majority of servers, it can be replicated on the state machine

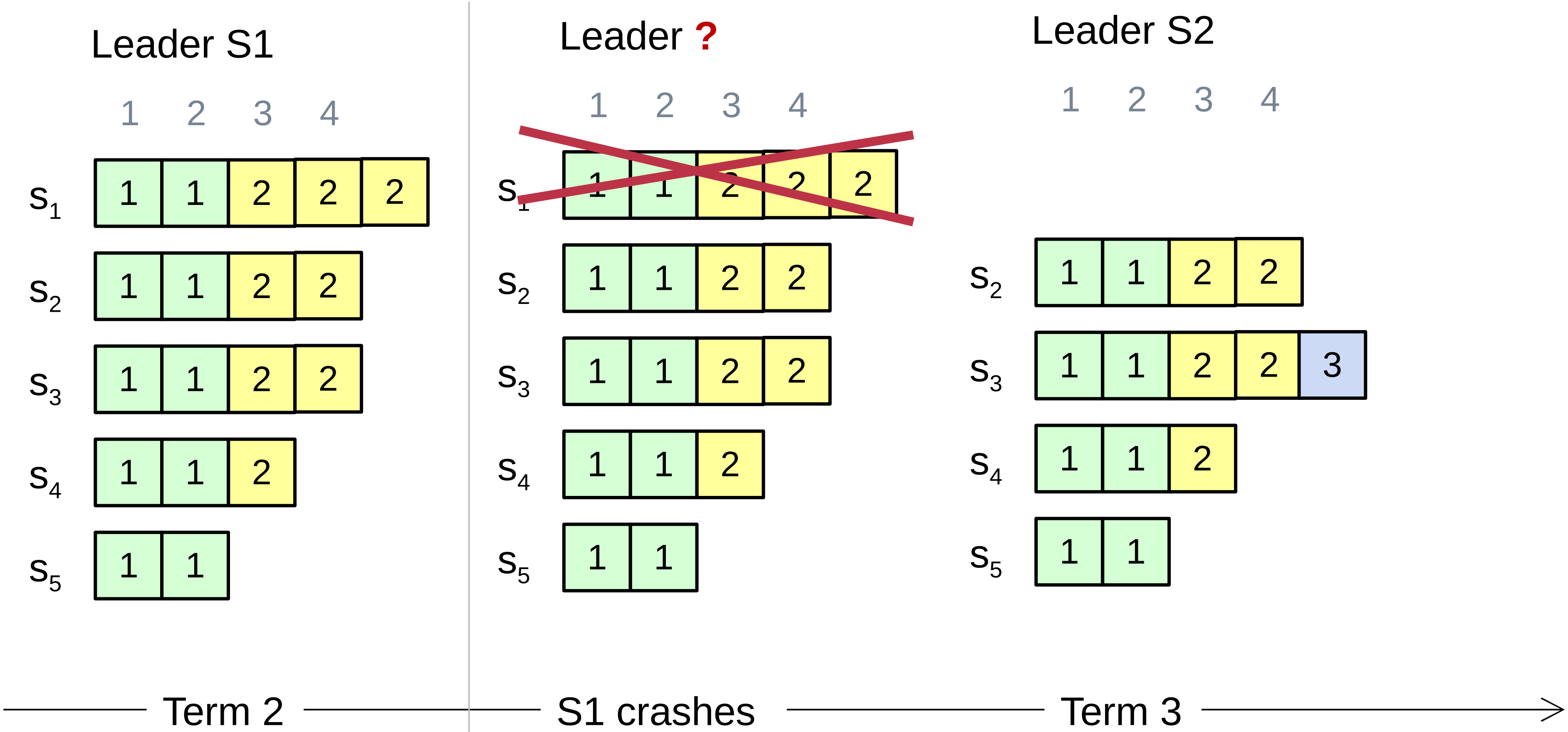
– Not always true on raft so far

Question: can index 4 be overwritten after appending to S3?



Case study: raft overwriting

Question: can index 4 be overwritten?



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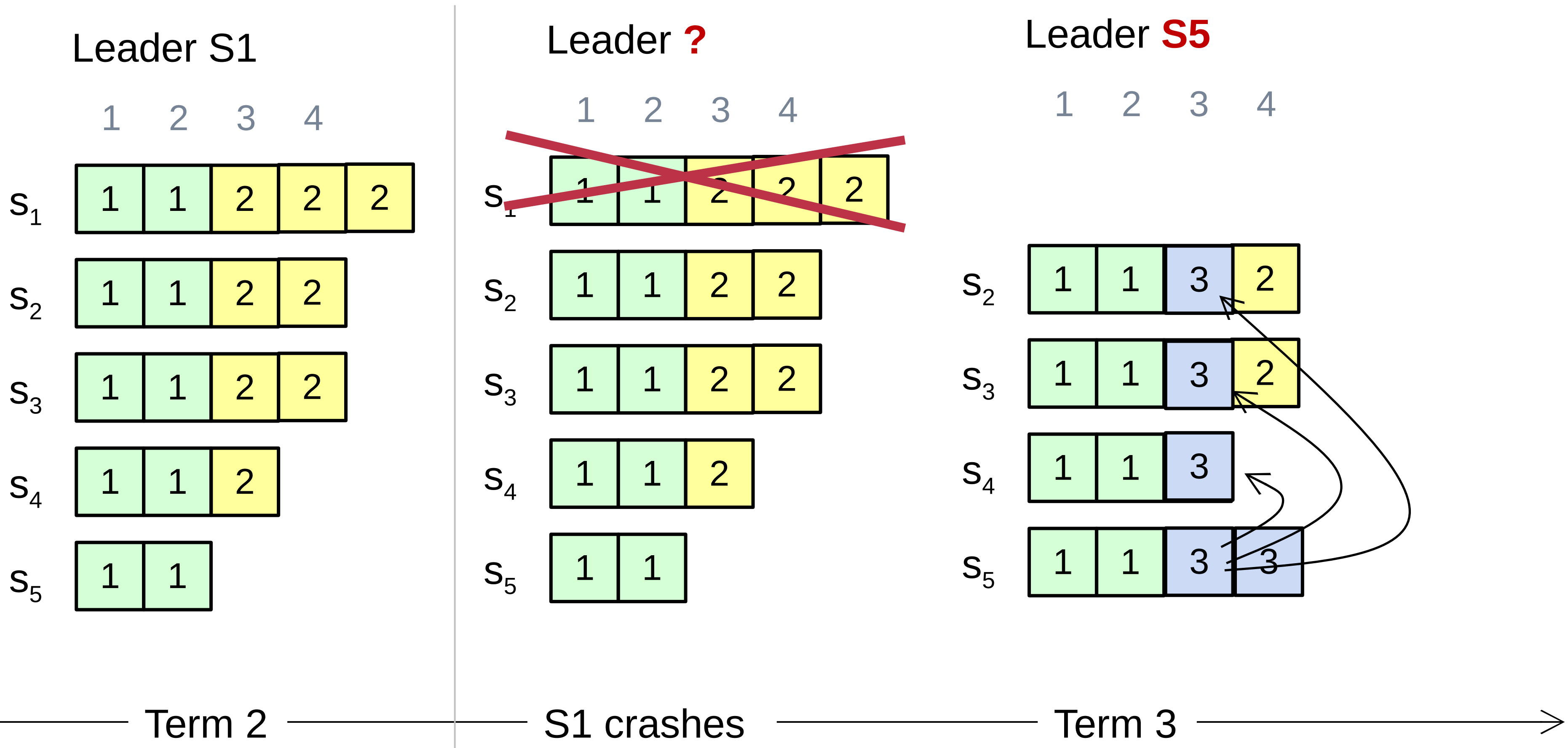


Safety requirement of the commit entry

Once a log entry has been applied to a state machine, no other state machine must apply a different value for that log entry

Raft safety property:

– If a leader has decided that a log entry is committed, that entry will be present in the logs of all future leaders (no overwritten)

This guarantees the safety requirement

– Leaders never overwrite entries in their logs

– Only entries in the leader's log can be committed

– Entries must be committed before applying to state machine

Committed → Present in future leaders' logs

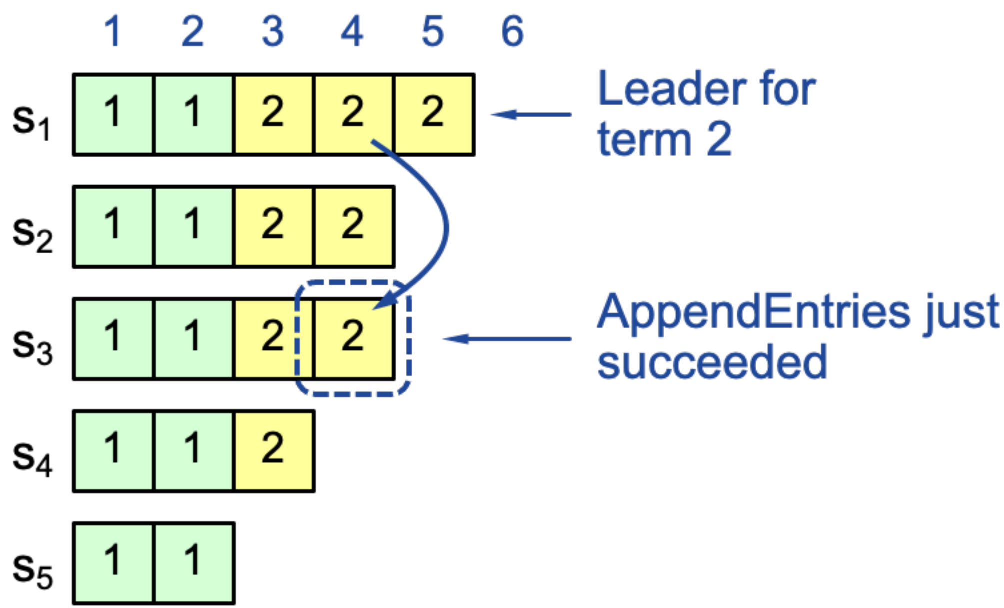
Restrictions on commitment

Restrictions on leader election

Goal: if a log entry has been replicated to majority followers, it is <u>likely</u> to commit

Committing Entry from the Current Term

Case #1/2: Leader decides entry in current term is committed



Question: how to prevent index 4 from being overwritten? Prevent S4 or S5 from becoming the leader

Picking the Best Leader

During elections, choose candidate with log most likely to contain all committed entries

- Candidates include log info in RequestVote RPCs
  (index & term of last log entry)
- Voting server V denies vote if its log is “more complete”:

```
(lastTermV > lastTermC) ||
(lastTermV == lastTermC) && (lastIndexV > lastIndexC)
```

- Leader will have “most complete” log among electing majority

Comparing the best leader

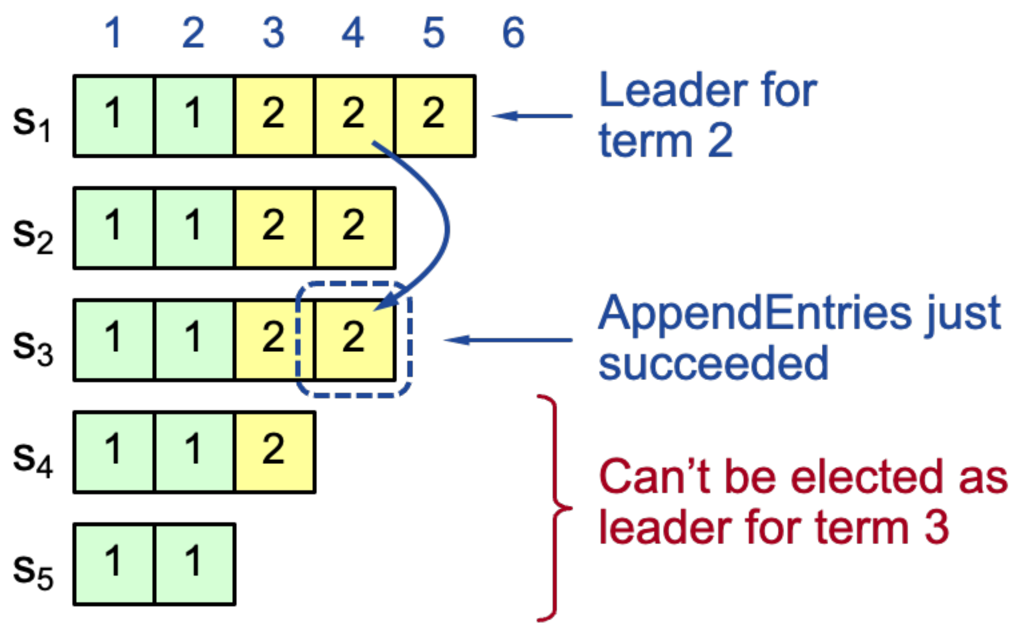
```
(lastTermV > lastTermC) ||
(lastTermV == lastTermC) && (lastIndexV > lastIndexC)
```

Question: which log is longer?

– S1 & S2, respectively

Committing Entry from the Current Term

Leader decides entry in current term is committed



Safe: leader for term 3 must contain entry 4

Complete picture of request vote RPC

Invoked by candidates to gather votes.

Arguments:

candidateId candidate requesting vote
term candidate's term
lastLogIndex index of candidate's last log entry
lastLogTerm term of candidate's last log entry

Results:

term currentTerm, for candidate to update itself
voteGranted true means candidate received vote

Implementation:

1. If term > currentTerm, currentTerm ← term
(step down if leader or candidate)
2. If term == currentTerm, votedFor is null or candidateId, and candidate's log is at least as complete as local log, grant vote and reset election timeout

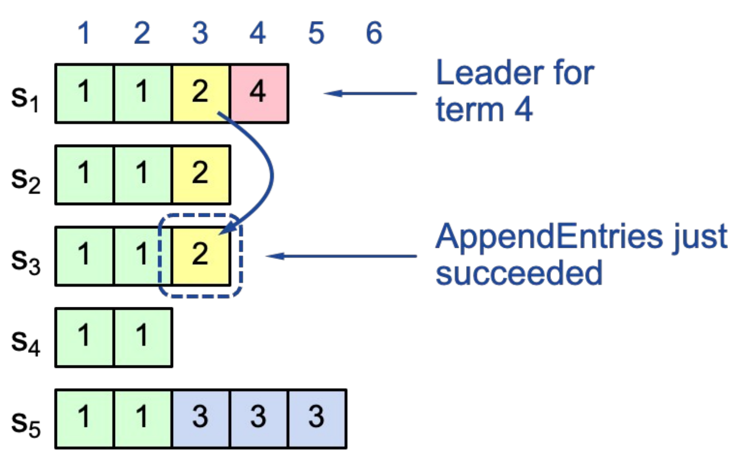
Commit rule for raft so far

If the log entry from the leader term is replicated to a majority of followers

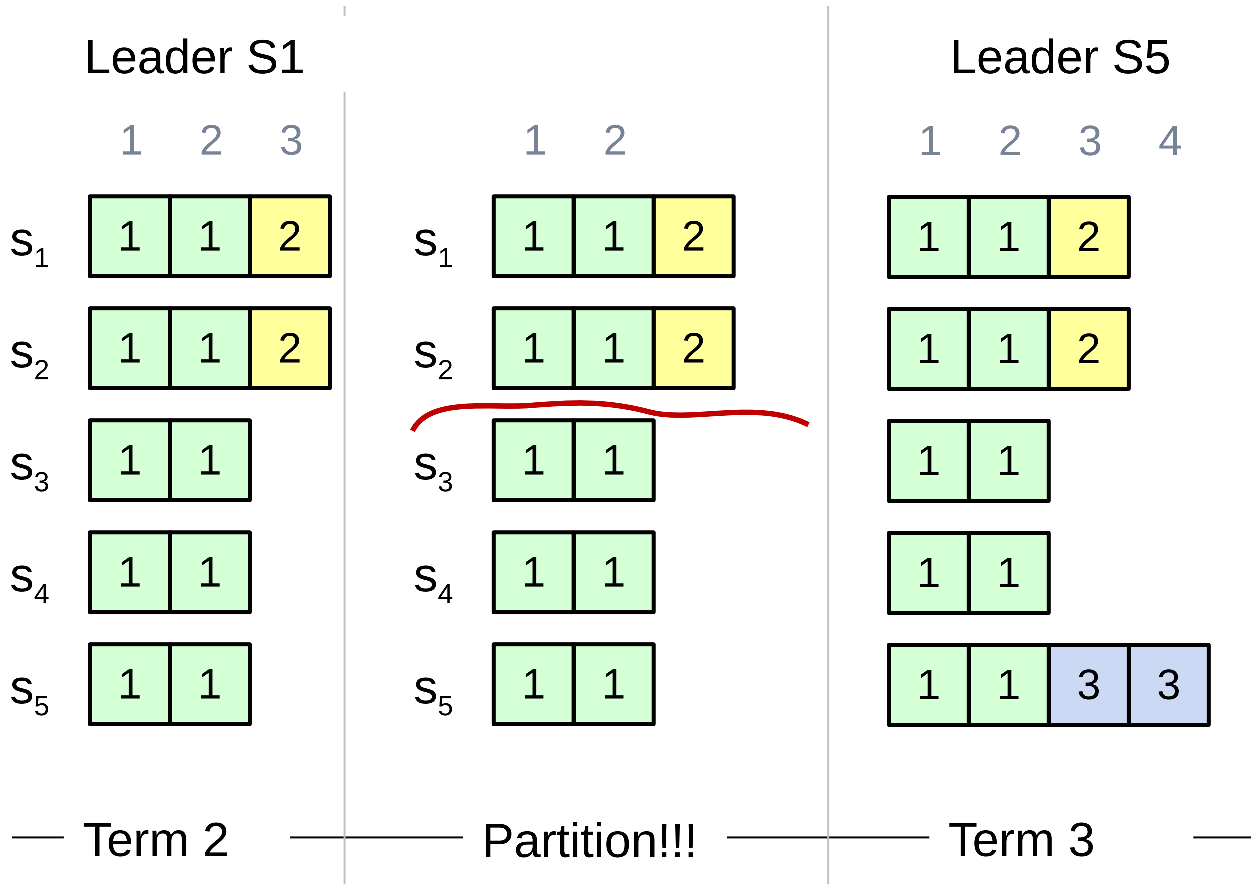
- Then we can treat it as committed
  - The later leader must contain the entry

But, what about replicating log entry from a previous term?

- The previous term’s entry may fail to reach a majority



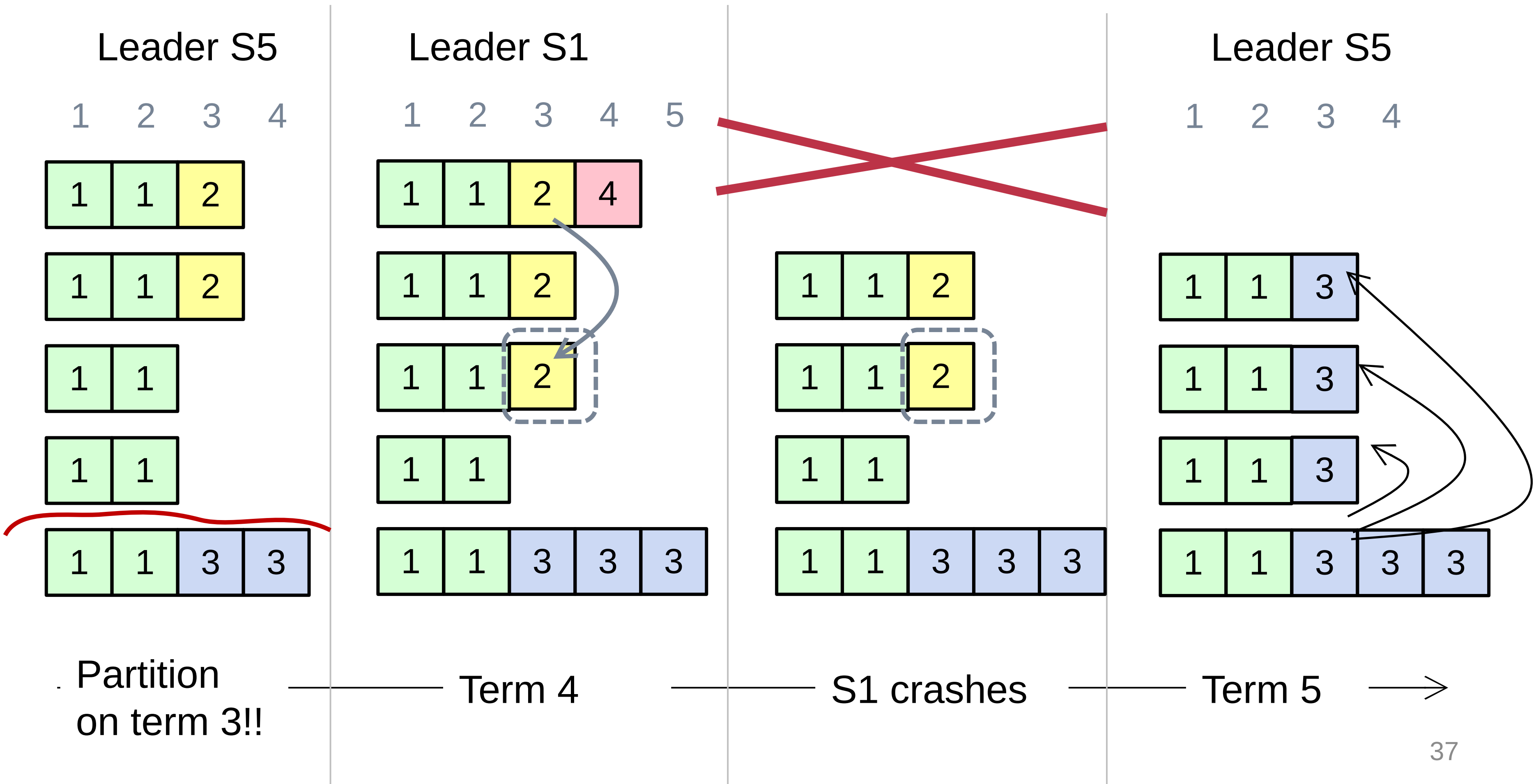
Case study: a majority replicated entry can be overwritten



Question: who can become the leader for term 3?

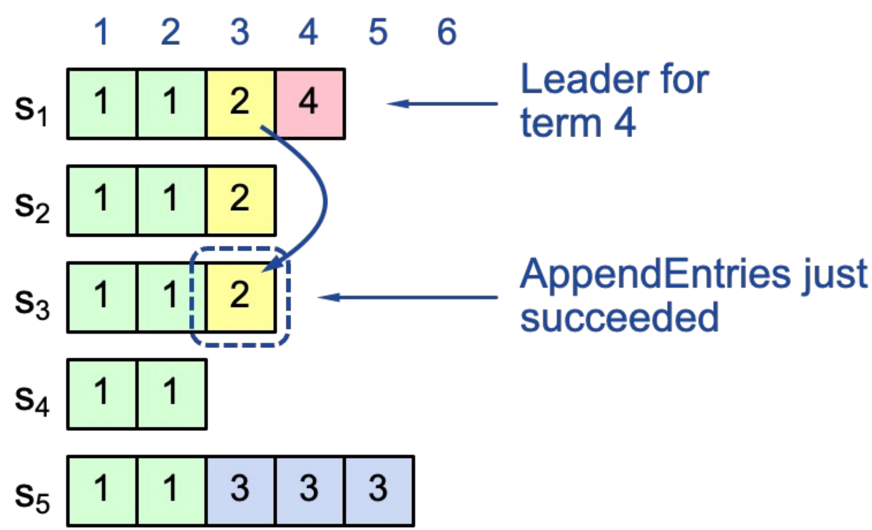
Partition again!!!

Case study: a majority replicated entry can be overwritten



Committing Entry from Earlier Term

Case #2/2: Leader is trying to finish committing entry from an earlier term



Entry 3 not safely committed:

- S5 can still be elected as leader for term 5
- If elected, it will overwrite entry 3 on S1, S2 and S3 (recall our previous example)

Commit rules

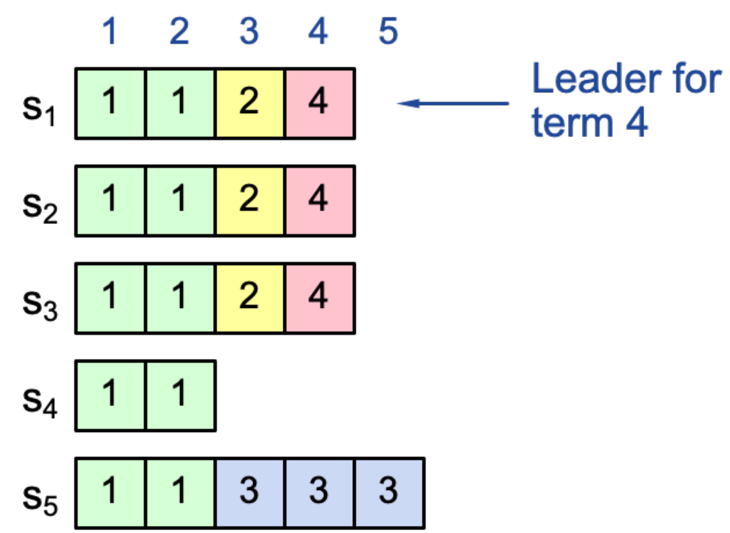
For a leader to decide an (previous) entry is committed:

- Must be stored on a majority of servers
- At least one new entry from leader's term must also be stored on majority of servers

This is because once entry 4 committed:

- s_5 cannot be elected leader for term 5
- Entries 3 and 4 both safe

Combination of election rules & commitment rules makes Raft safe



Detailed distributed database study:

Google Spanner

Problems faced by large-scale company

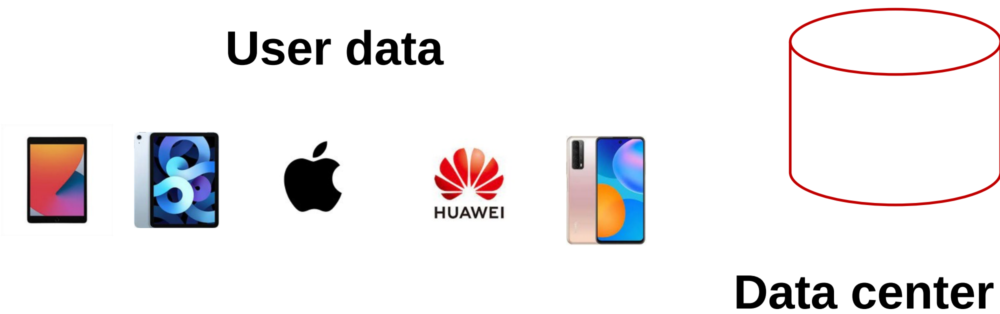
Large-scale dataset

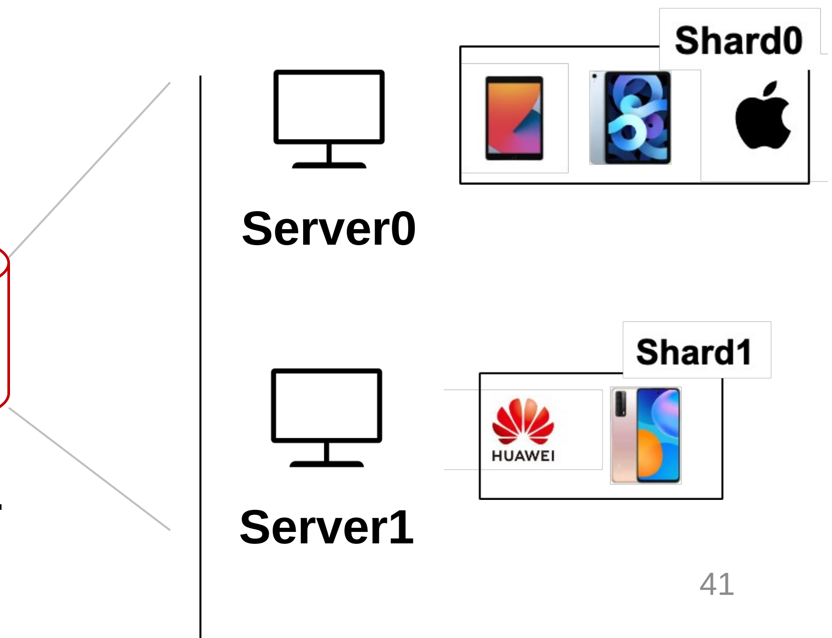
– **Facebook** has more than **1 billion** of images uploaded **weekly**, **Baidu** stores **tens of billions** of web pages

Question: how to store such a large dataset?

Solution: sharding

– Each shard is stored in an LSM storage





How to shard?

A topic orthogonal to the design of Spanner

Possible way:

- Hashing, node id = hash(key) % num_of_nodes
- Range partitioning
  - E.g., coordinators record that 0-100 keys are at shard 0, etc.

A good partition need to consider many factors, e.g.,

- Reduce cross shard transaction (no 2PC for multi-site transactions)
  - Why? Shards are likely to span across two machines, which needs two-phase-commit for execution

We will not focus sharding in this lecture

Problems faced by large-scale company

Fault tolerance

- Large-scale companies are built over **large-scale datacenters** (>10K servers)
- Machine failures are particular common in large-scale datacenters

> “Suppose a cluster has ultra-reliable server nodes with a stellar mean time between failures (MTBF) of 30 years (10,000 days)—a cluster of 10,000 servers will see an average of one server failure per day. ”

Problems faced by large-scale company

Fault tolerance

- Large-scale companies are built over large-scale datacenters (>10K servers)
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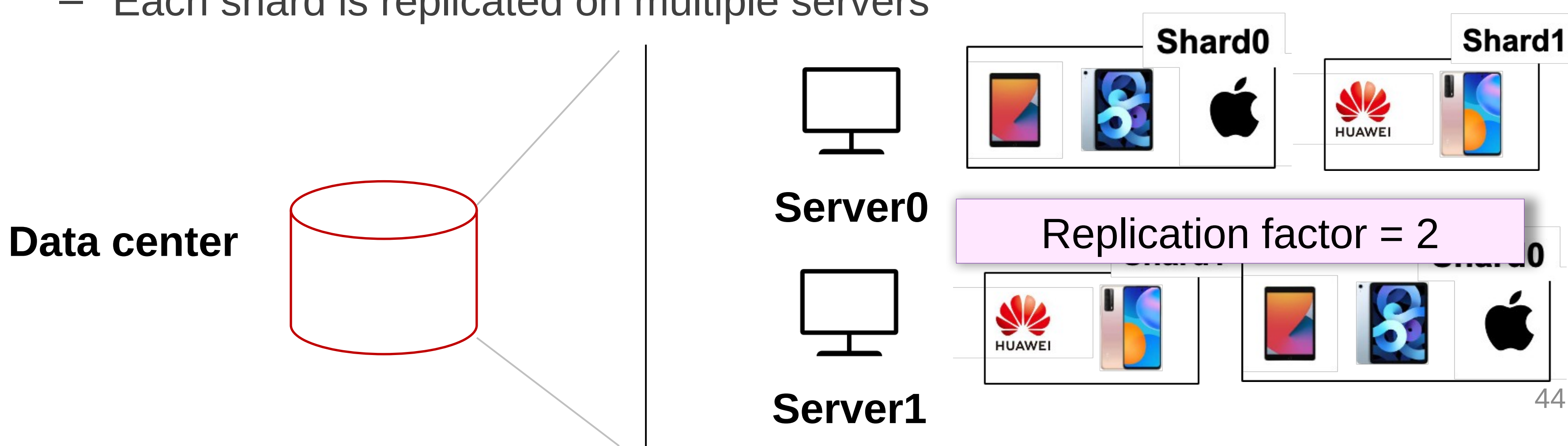
Solution: replication

- Each shard is replicated on multiple servers

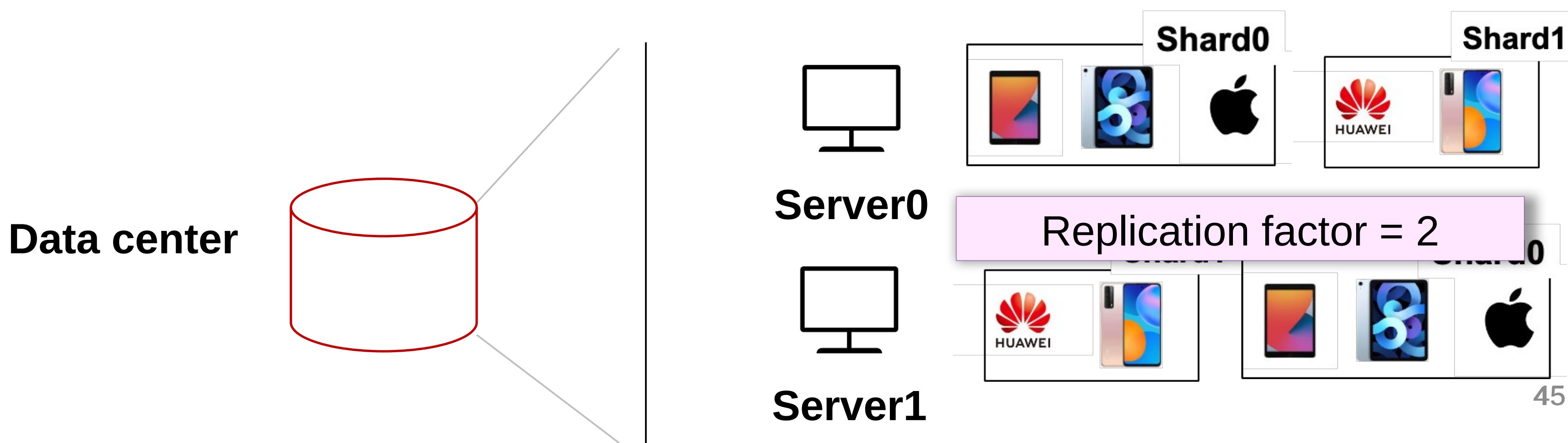


Question: is a replication factor of 2 ok?

Ok means can achieve single-copy consistency

Whether OK depends on the scenarios (& setup)

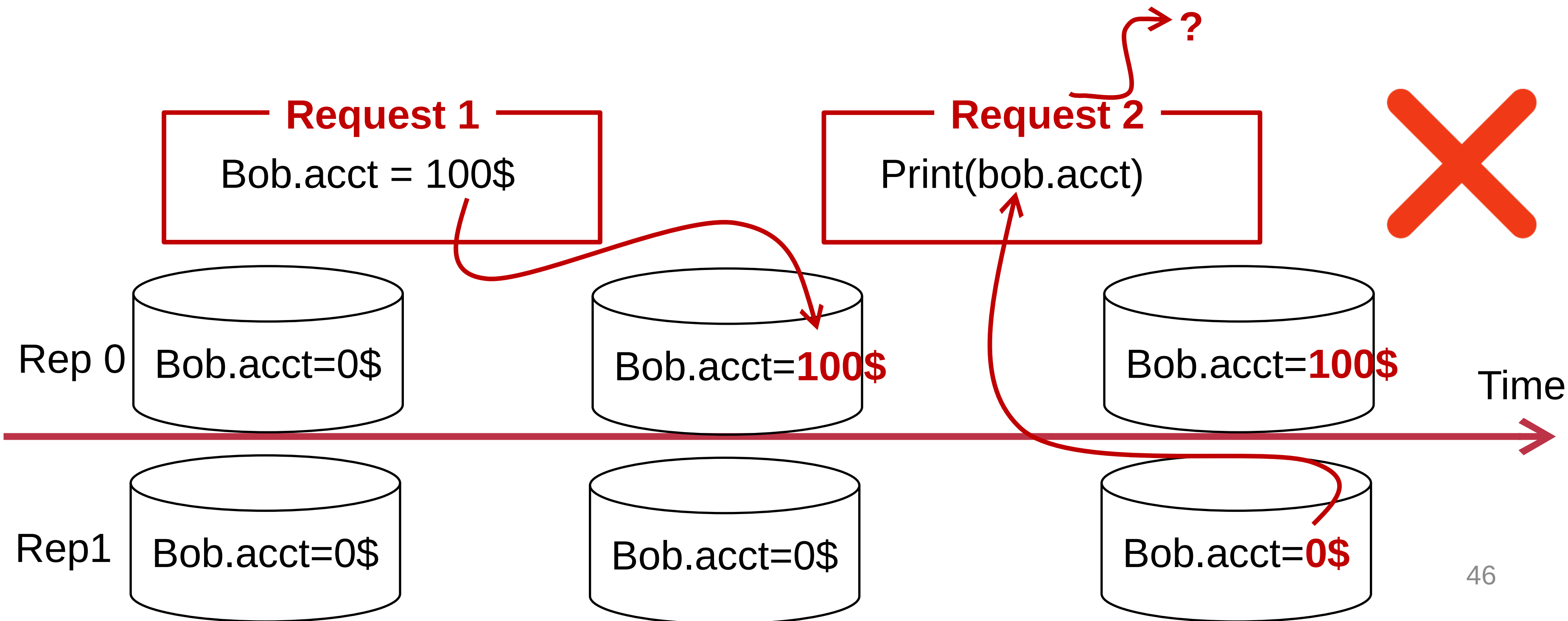
- Under primary-backup replication, it is ok
    - We have a view server
- Under Paxos & Raft, it's not! (no majority)



Replications causes consistency problem

Could not achieve single-copy consistency if allow client arbitrary reads/writes

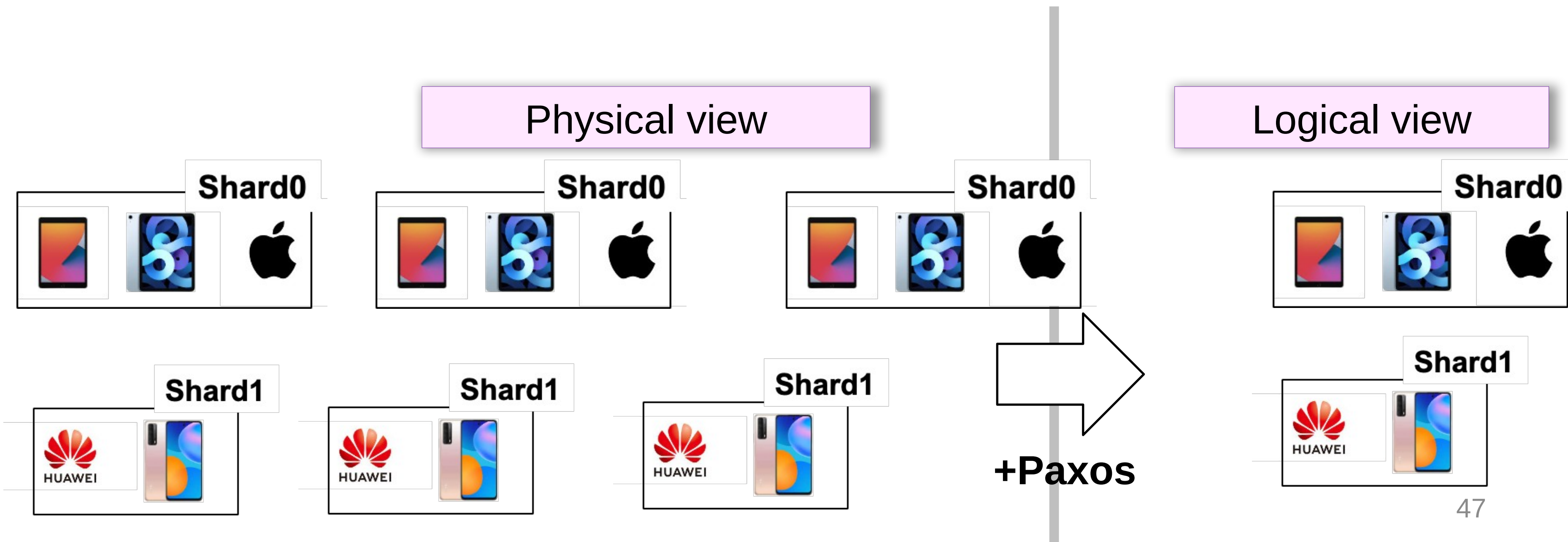
– E.g., consistency among replicas



Solution: Paxos for single-copy consistency

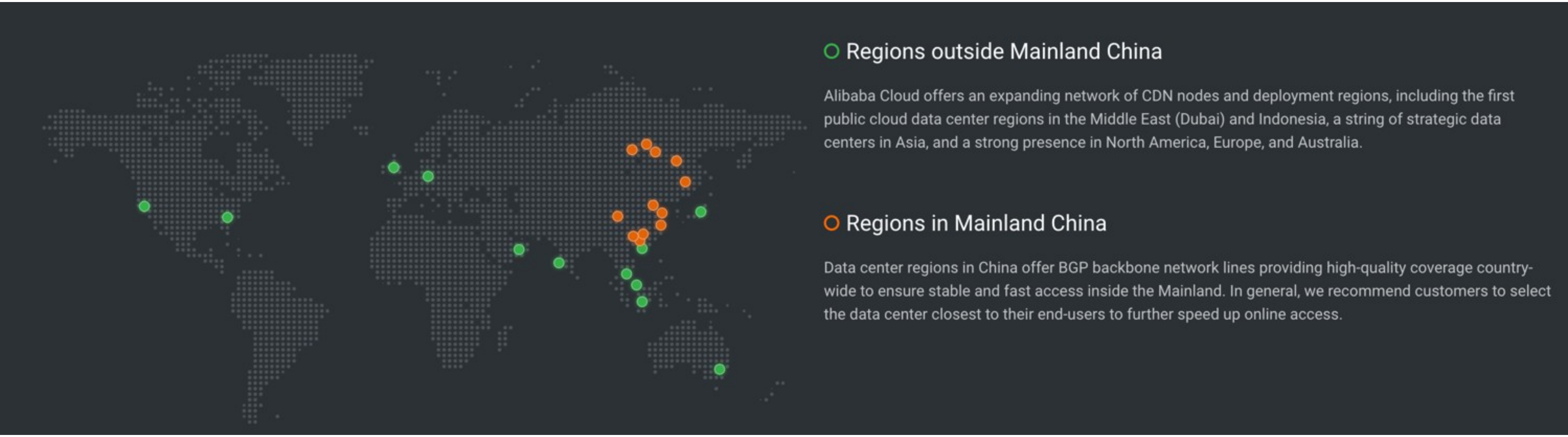
Spanner uses Multi-Paxos w/ leaders to abstract the replicated shard

– On the developer's perspective, a shard w/ replication can be viewed as a single-shard (single-copy consistency)



Geo-replicated datacenters

e.g., The **locations** of Alibaba datacenters



What can go wrong if we only replicate the data **only within one datacenter**?

Source: https://www.alibabacloud.com/zh/global-locations

Replication within a DC is insufficient

A single datacenter (DC) can fail due to natural disasters

A DC cannot serve requests well (in low latency) across the planet

DC in the US

Users from china will suffer from long latency

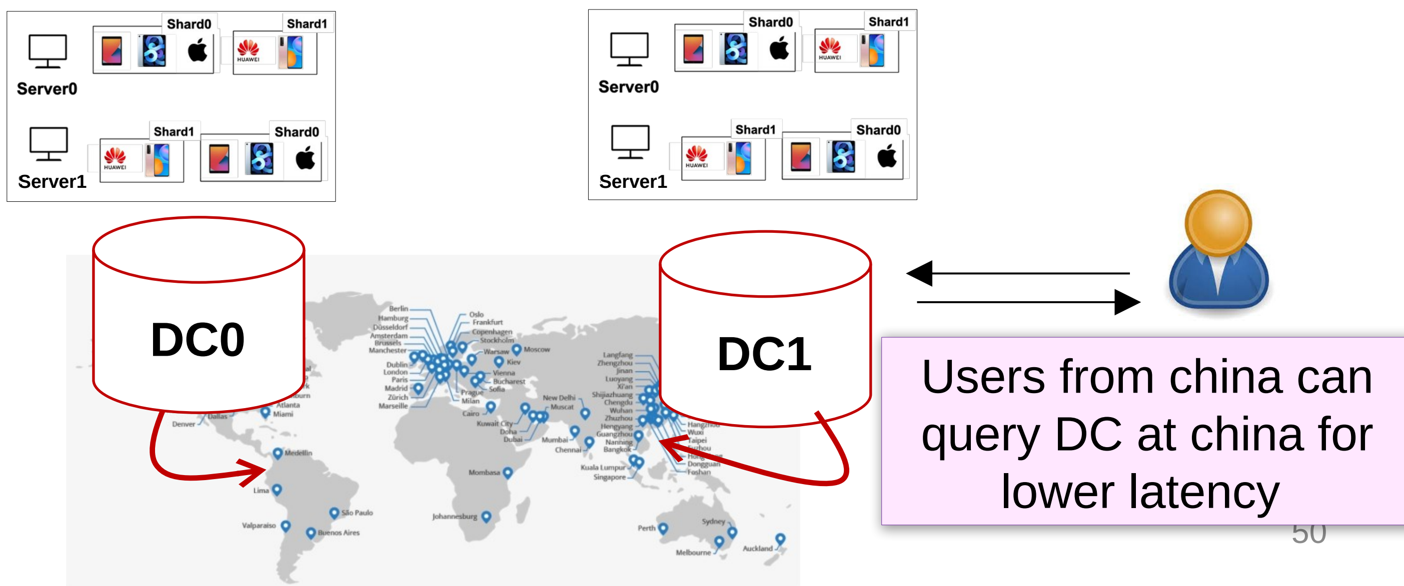
Spanner further replicates the data across datacenters

Data sharded & replicated over many machines & **datacenters**

- Tolerate machine failures across datacenters & lower request latency

User data

Data centers



2-phase commit + 2-phase locking

With Paxos, multiple replicas of a shard can be viewed as one

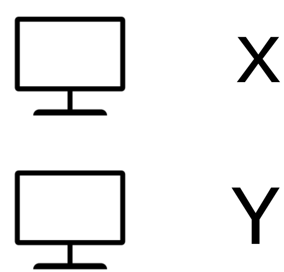
- E.g., the developer sees an X, but X is actually replicated multiple times

However, Paxos is insufficient

- A user request (TX) can touch multiple shards
- How to coordinate these accesses?

Using two-phase locking for before-or-after atomicity & two-phase commit for multi-site atomicity

```
tx.begin()
x = x + 1
y = y + 1
tx.end()
```



Big picture of Spanner

How to scale to a larger dataset?

- Shading

How to tolerate failures across machines & datacenters?

- Replicated state machine w/ Paxos
    - Can also use raft (e.g., TiDB)

How to execute read-write transaction to ensure before-or-after atomicity and multi-site atomicity?

- 2PL + 2PC

Execution flow of read-write transaction(TX)

Read-write TX:

– A TX that both reads and writes the data

```
tx.begin()
x = x + 1
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tx.end()
```

Client

Execution flow of read-write transaction(TX)

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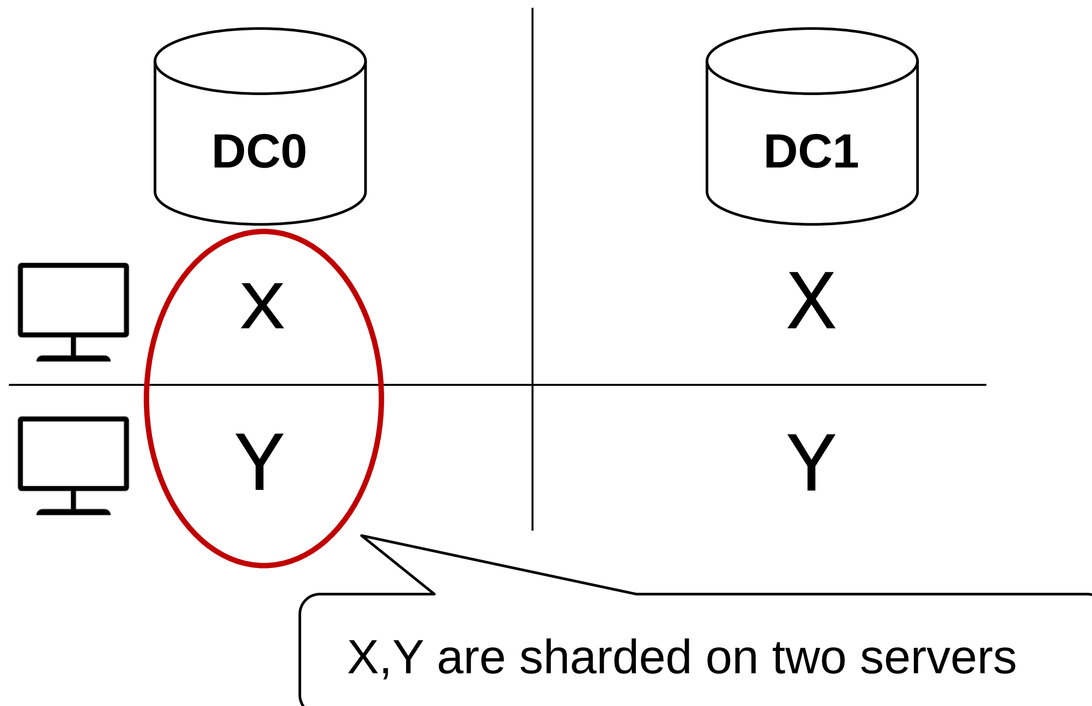

Execution flow of read-write transaction(TX)

Each shard is replicated

– The run PAXOS to behavior as a single (logical) shard

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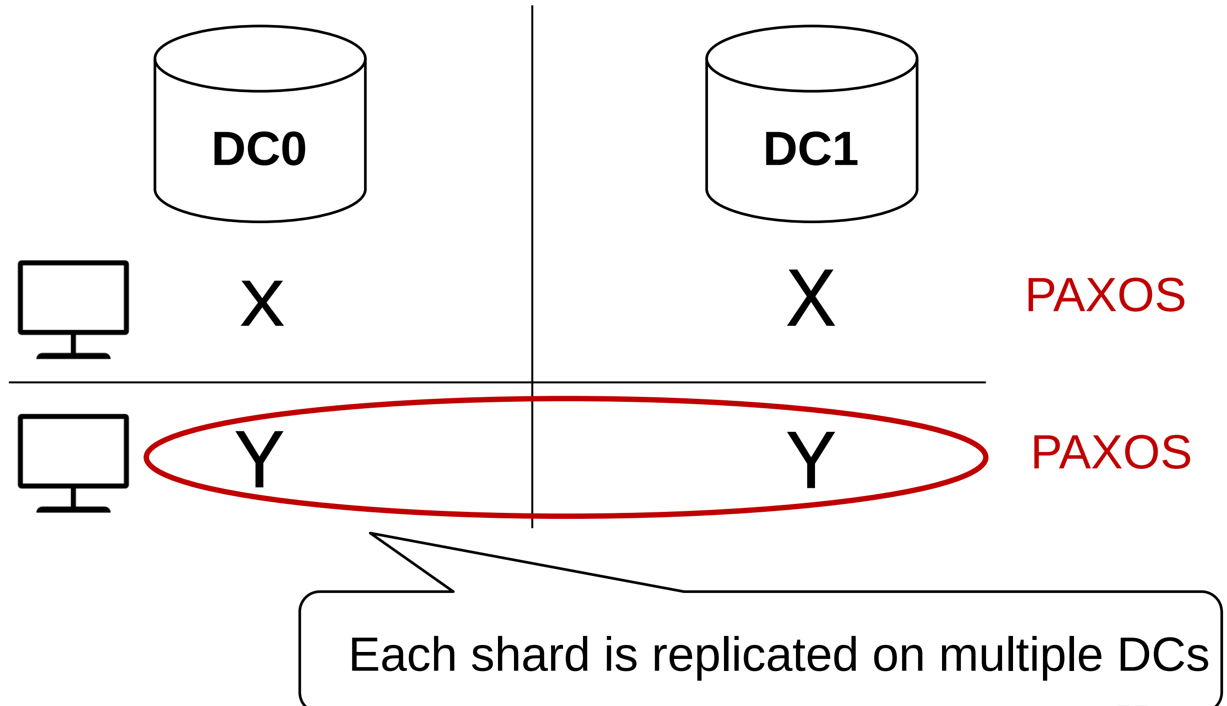


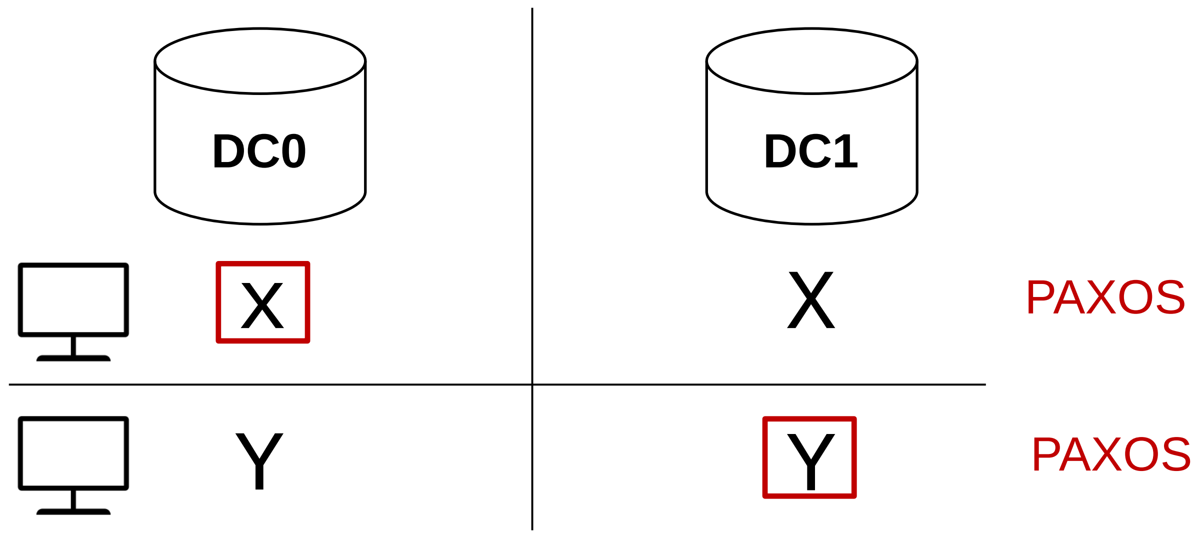
Execution flow of read-write transaction(TX)

Each shard is replicated w/ PAXOS

– A replica is selected leader to simplify execution (no Raft introduced then)

```
tx.begin()
x = x + 1
y = y + 1
tx.end()
```

Client



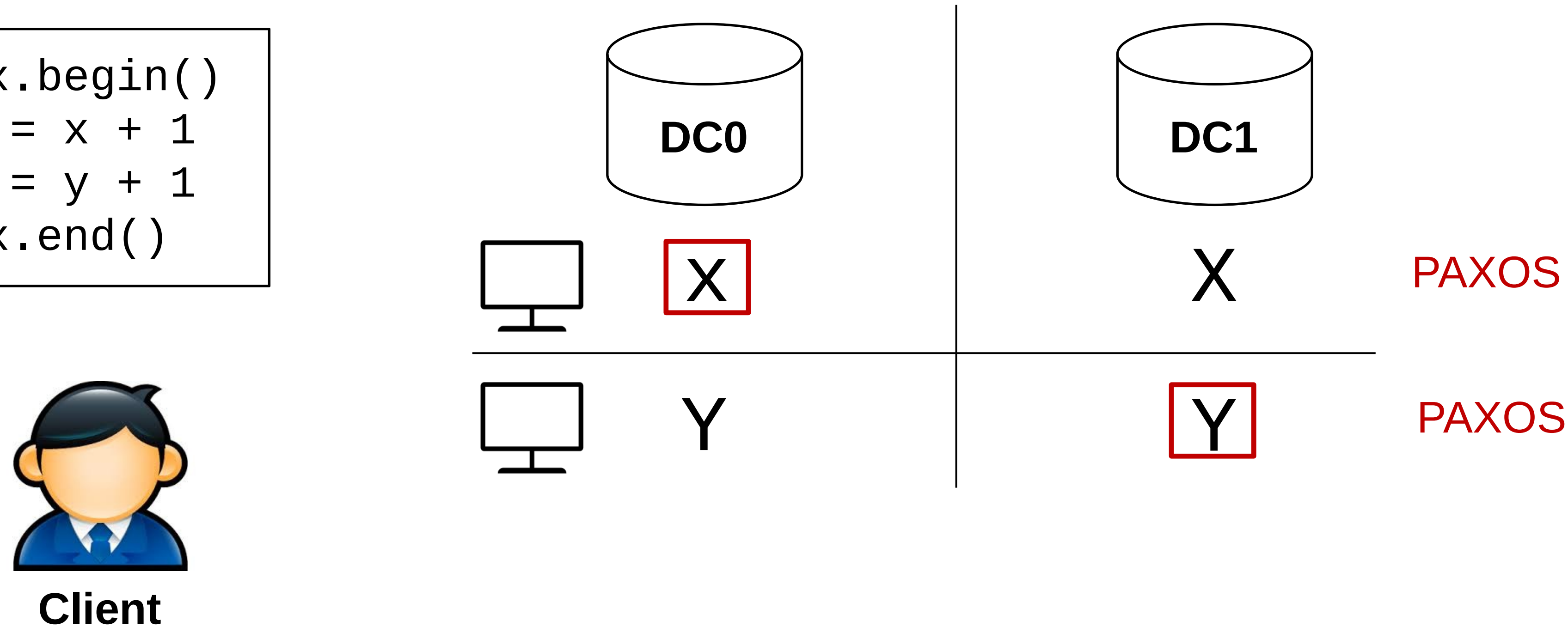
Leader will execute request
Start a PAXOS proposal accordingly

Execution flow of read-write transaction(TX)

Spanner uses standard 2PC & 2PL for executing read-write TX

– Two-phase commit & two-phase locking

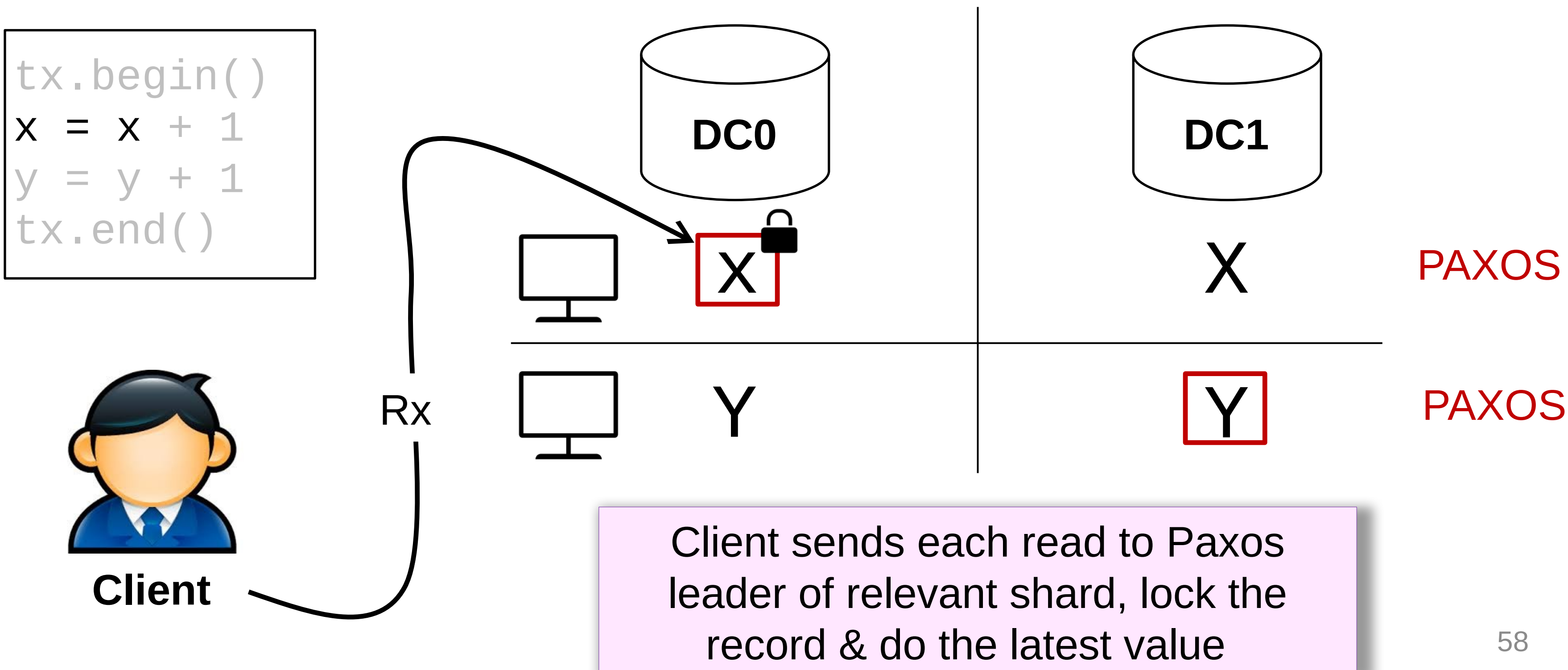
```
tx.begin()
x = x + 1
y = y + 1
tx.end()
```



Execution flow of read-write transaction(TX)

Upon read, the leader will return the latest value of data

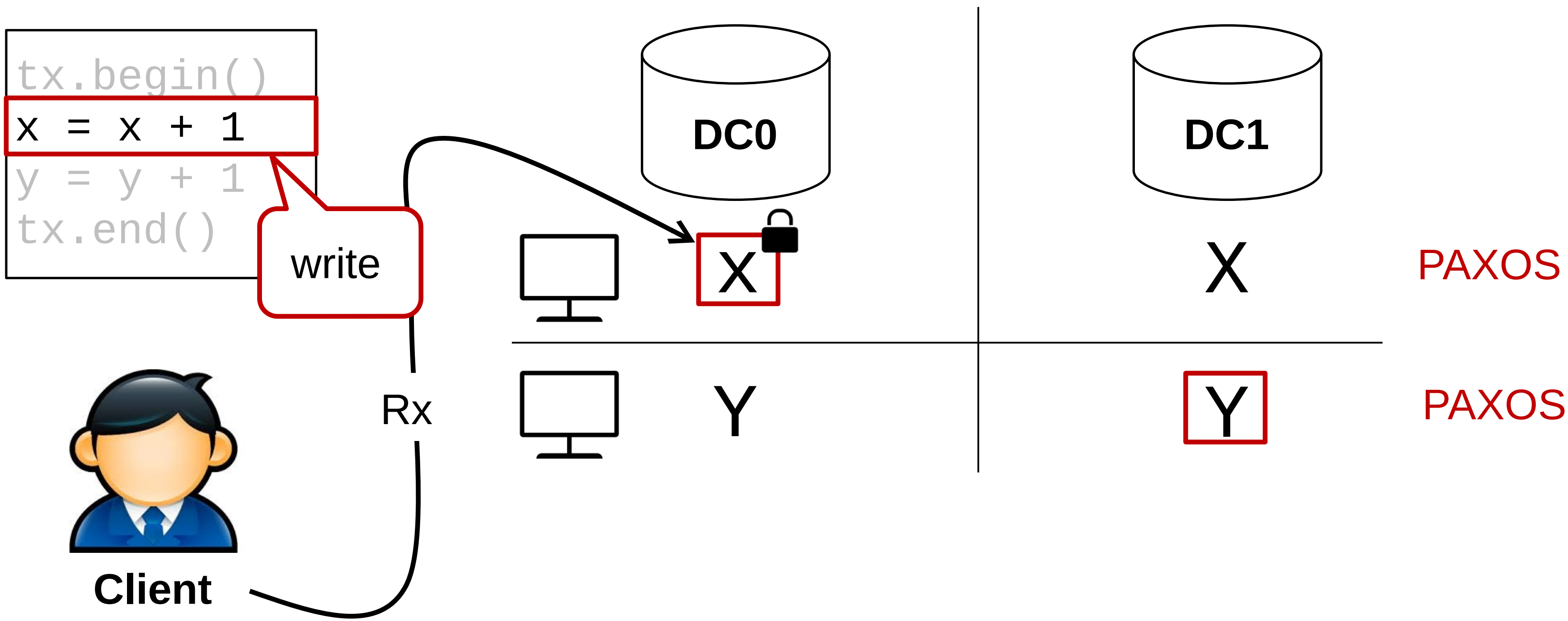
– Also hold the lock on it (2PL)



Execution flow of read-write transaction(TX)

The write will be buffered at local client

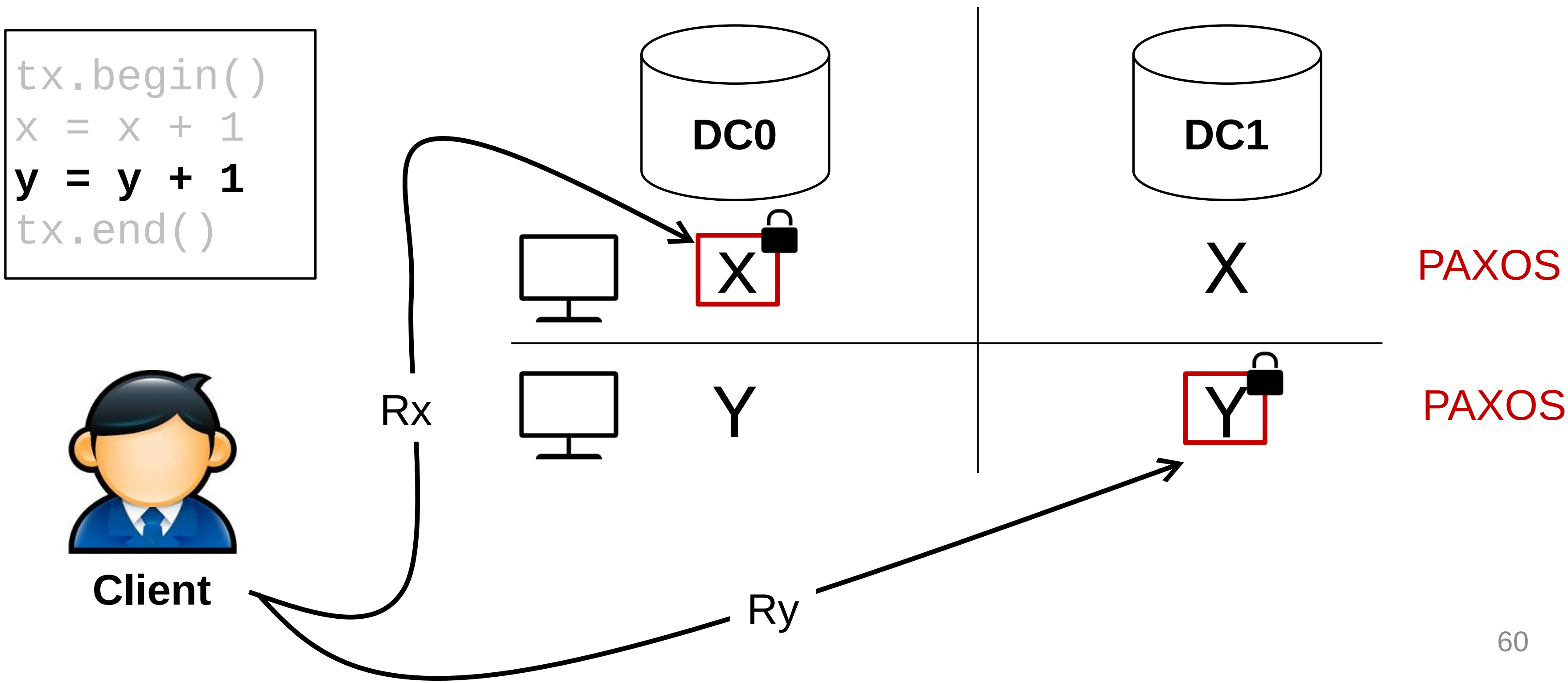
– They will be written back via 2PC at the TX's commit time



Execution flow of read-write transaction(TX)

Commit will use 2PC to write all the updates back

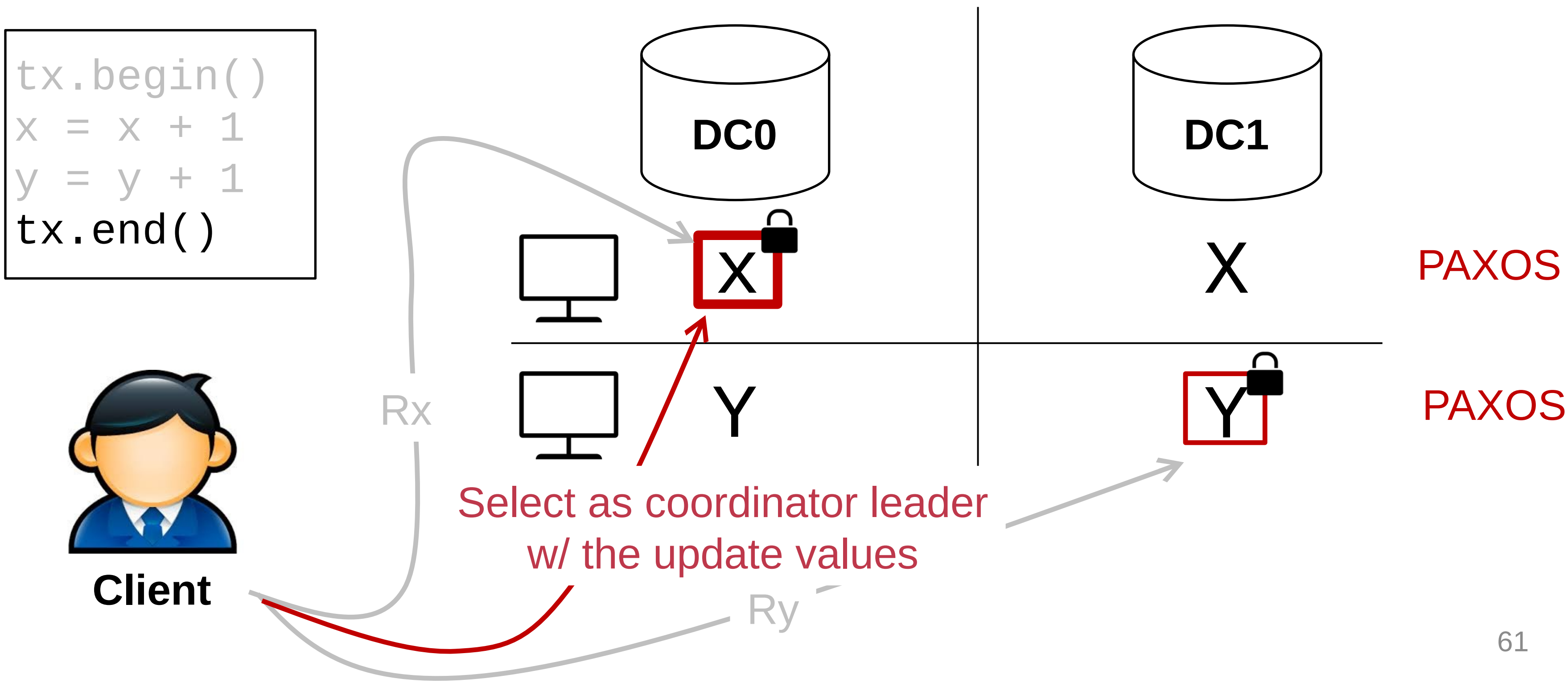
– It will choose a shard leader as the coordinator leader



Execution flow of read-write transaction(TX)

The coordinator uses its PAXOS group to replicate the TX states

– Why? Otherwise, the 2PC may fail if the coordinator crashes!

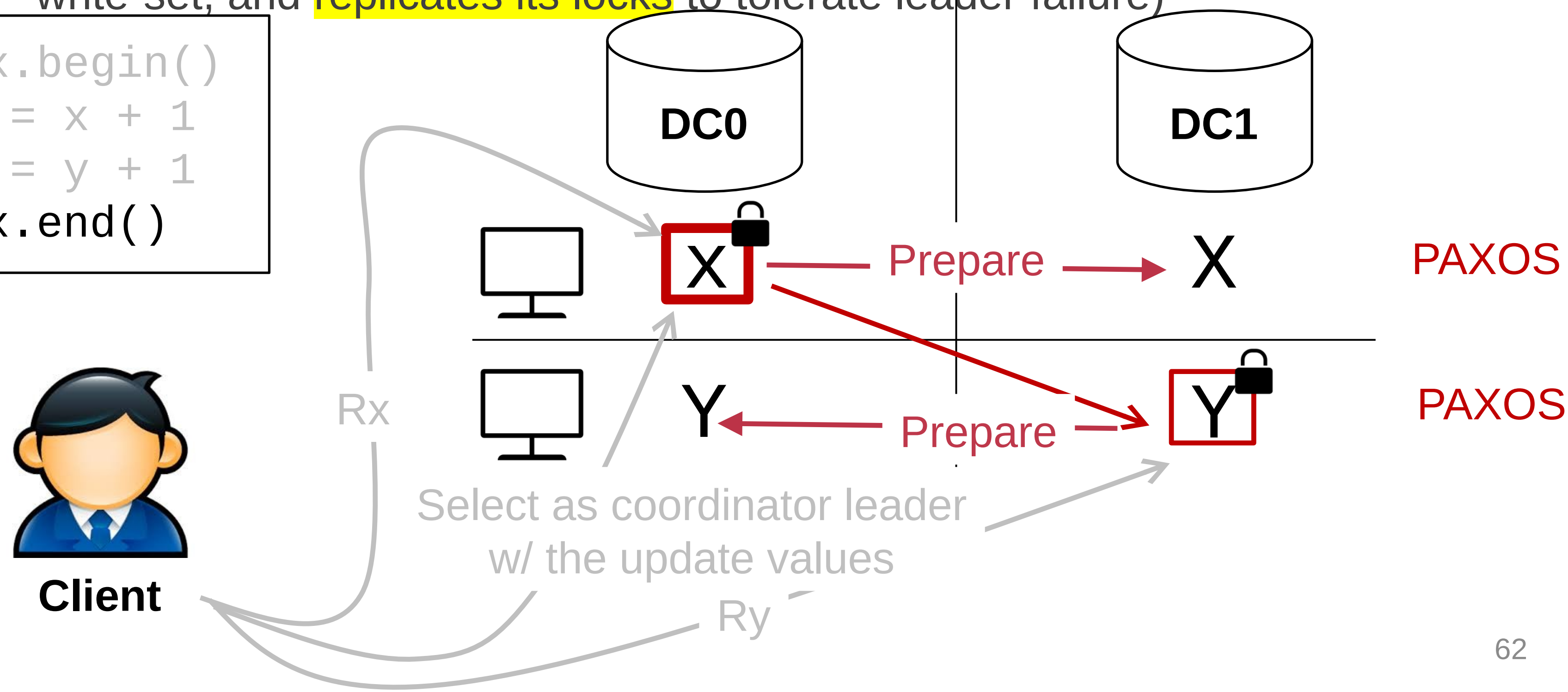


Execution flow of read-write transaction(TX)

The coordinator leader will send the update values to the host shard leader

– And each shard leader execute the **prepare** with PAXOS (including locking the write-set, and replicates its locks to tolerate leader failure)

```
tx.begin()
x = x + 1
y = y + 1
tx.end()
```

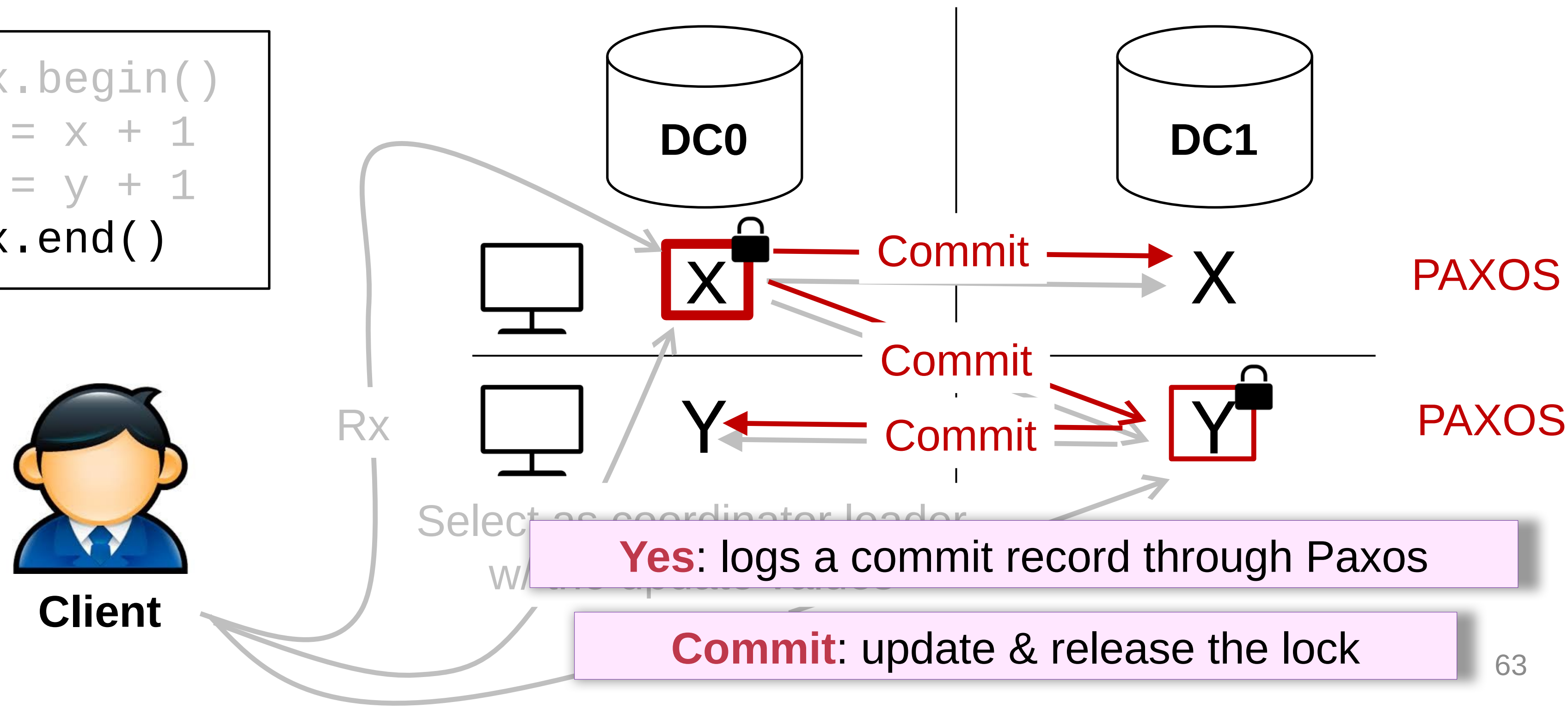


Execution flow of read-write transaction(TX)

If all shard leader returns the Yes to the coordinator leader

– The coordinator log the commit decision (via PAXOS) & commit others

```
tx.begin()
x = x + 1
y = y + 1
tx.end()
```



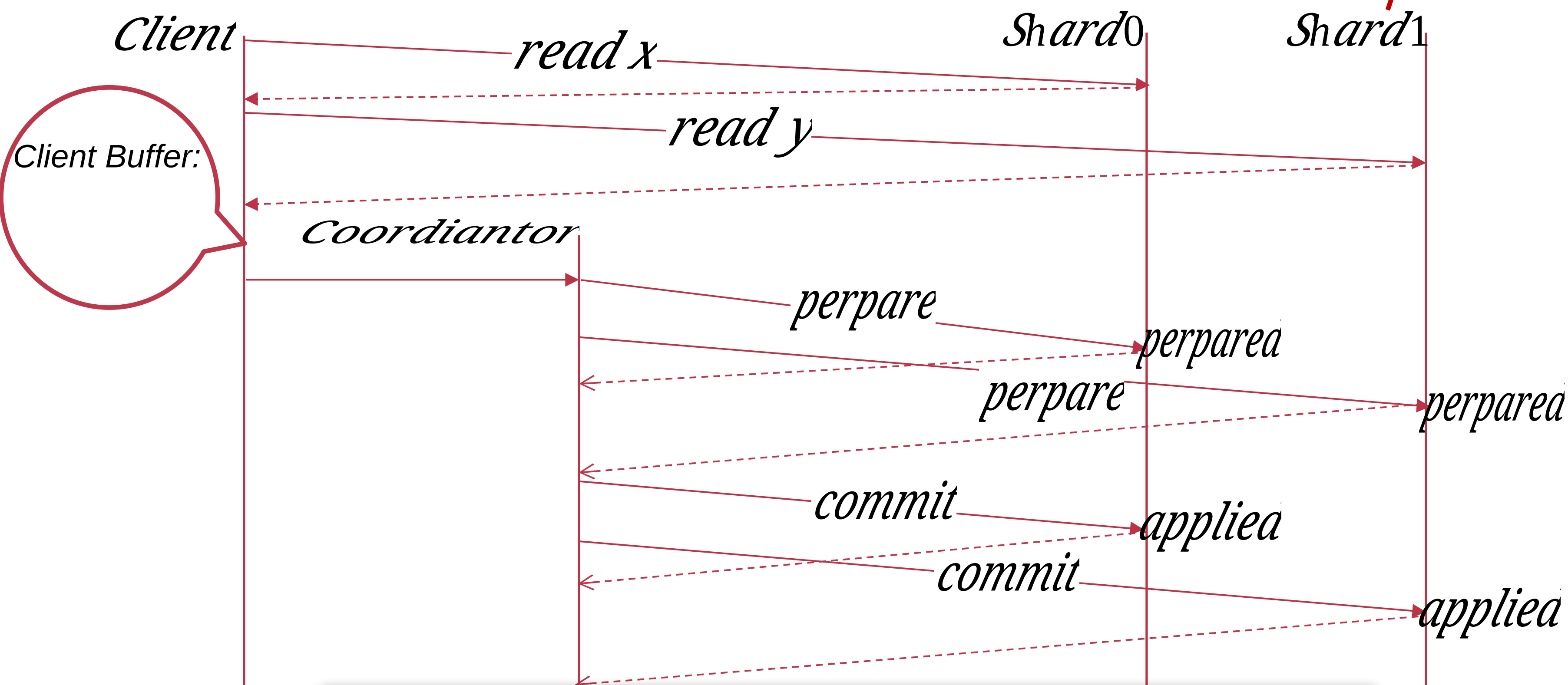
Yes: logs a commit record through Paxos

Commit: update & release the lock

Read-write TX: put it all together

Multiple physical nodes backed by PAXOS

$2PL + 2PC + Paxos$



Note: we skip the network RTTs of Paxos

Read-write TX: put it together

Spanner’s read-write TX gives a strong abstraction to the user

- A single-thread “machine” that never fails (even tolerate natural disasters)
- The “machine” has “unlimited” storage capacity
    - As long as the machines in the datacenters have enough capacity to store the data

But, what are the costs?

Read-write TX of Spanner so far

Techniques: 2PL & 2PC & Paxos

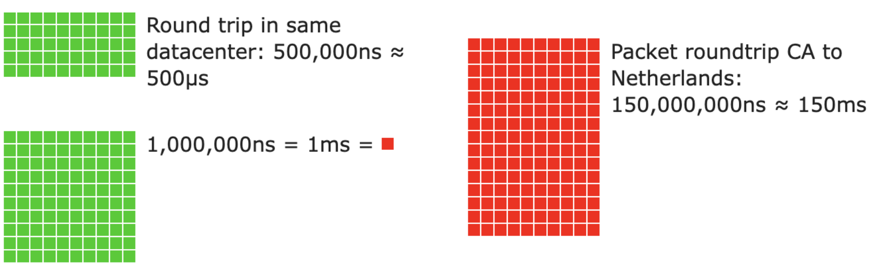
– Not optimized– Spanner finds writes are infrequent at Google

– Maybe slow: many messages sent between machines

- Also, across datacenters!

Table IV. Two-Phase Commit Scalability. Mean and Standard Deviations over 10 Runs

participants	latency (ms)	
	mean	99th percentile
1	14.6 ± 0.2	26.550 ± 6.2
2	20.7 ± 0.4	31.958 ± 4.1
5	23.9 ± 2.2	46.428 ± 8.0
10	22.8 ± 0.4	45.931 ± 4.2
25	26.0 ± 0.6	52.509 ± 4.3
50	33.8 ± 0.6	62.420 ± 7.1
100	55.9 ± 1.2	88.859 ± 5.9
200	122.5 ± 6.7	206.443 ± 15.8

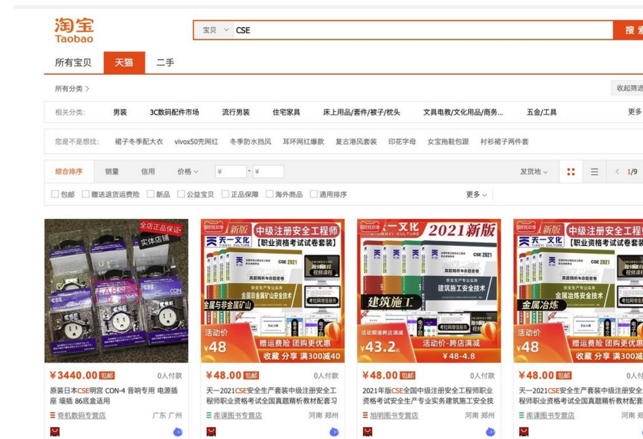


Source: https://colin-scott.github.io/personal_website/research/interactive_latency.ht

What about read-only TX?

Problem: read-only TXs are massive faced by Google

- Each read-only TX may touch many datasets



Drawbacks of 2PL (in read-only TX)

- Acquire the read lock is costly (need Paxos!)
- May possibly lock the items very long

Solution

- Multi-version concurrency control

Review: Multi-version Concurrency Control

Each data item has multiple versions

– When accessing different versions of data, probably no conflict!

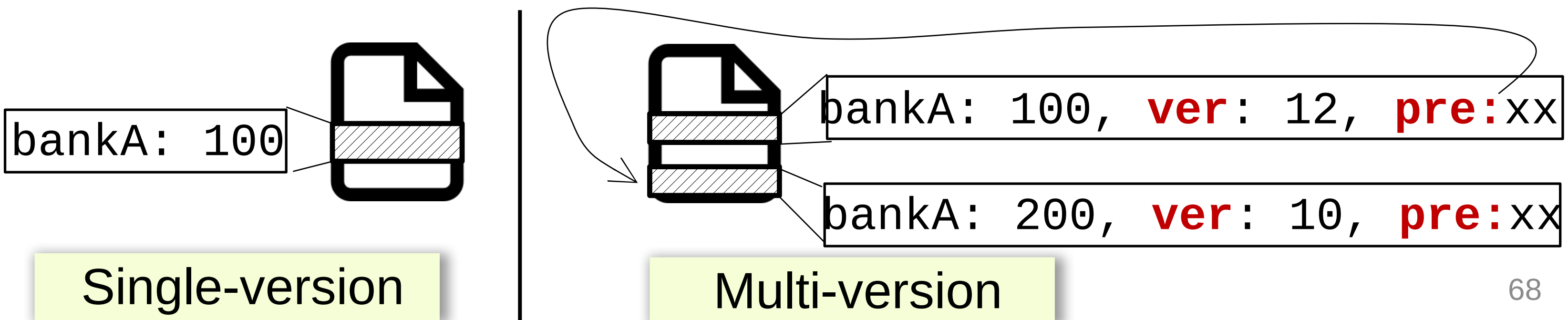
Key (high-level) idea

– Writes **don't overwrite** the original data

– Instead, writes **install new versions** of data

– Reads read from a "**snapshot**" of data

Benefits: no lock or validation during TX's execution



Review: Snapshot Isolation

A popular multi-version concurrency control (MVCC) scheme

- Transactions will get start and commit timestamp
- Use start timestamp to find the snapshot to read
- Use commit timestamp to install new versions

Transactions:

- **WRITEs** a local **buffer** (similar to OCC)
- **READs** a "**snapshot**" of entire data image
- **COMMIT** only if no **write-write** conflict
    - Install new versions of data

Drawback: no serializability guarantee !

Idea: use 2PL for read-write TX, MVCC for read-only TX

Still use 2PL for read-write TX

- The execution of TXs is always serializable!

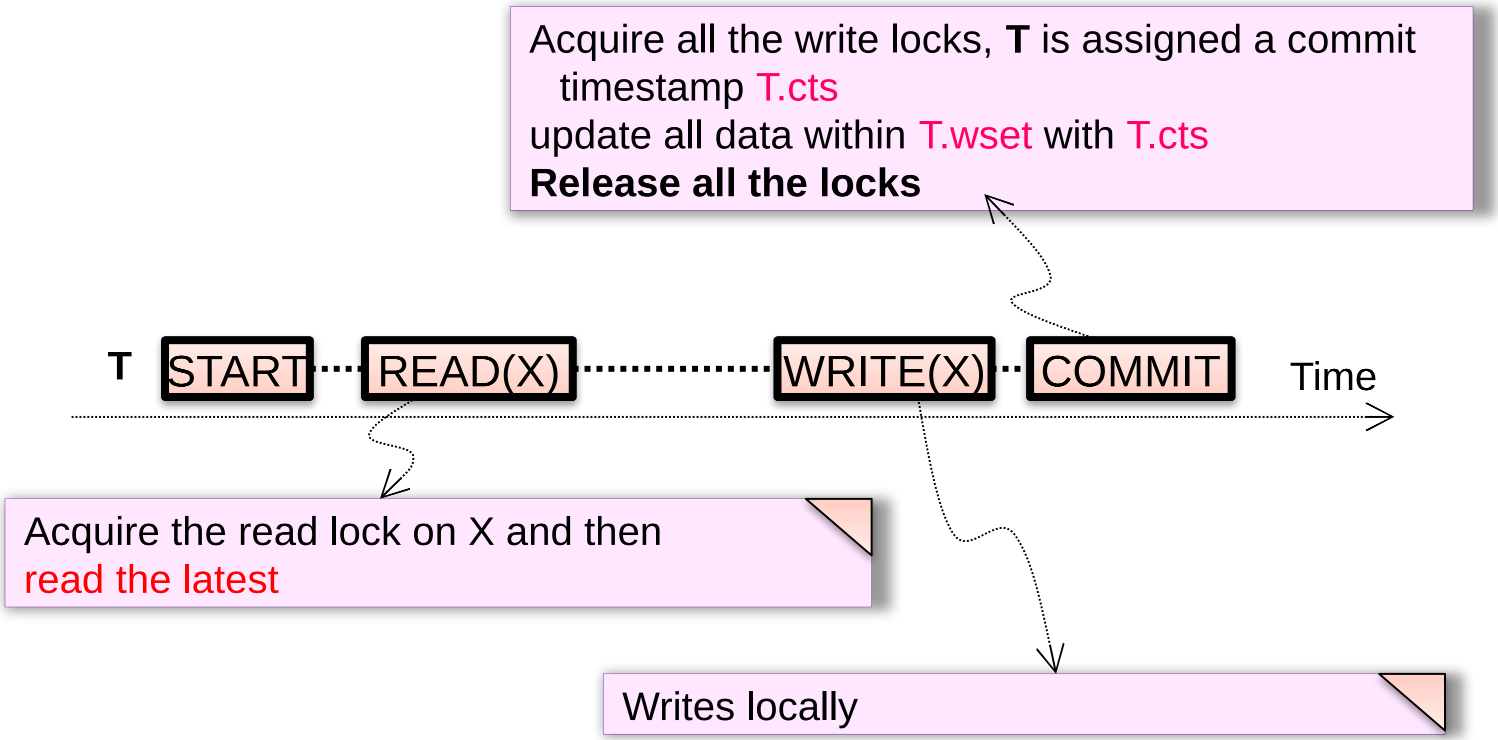
Use MVCC to execute the read-only TX

- Read-only TX no-longer to acquire locks for the execution

Note that similar technique also applies to OCC

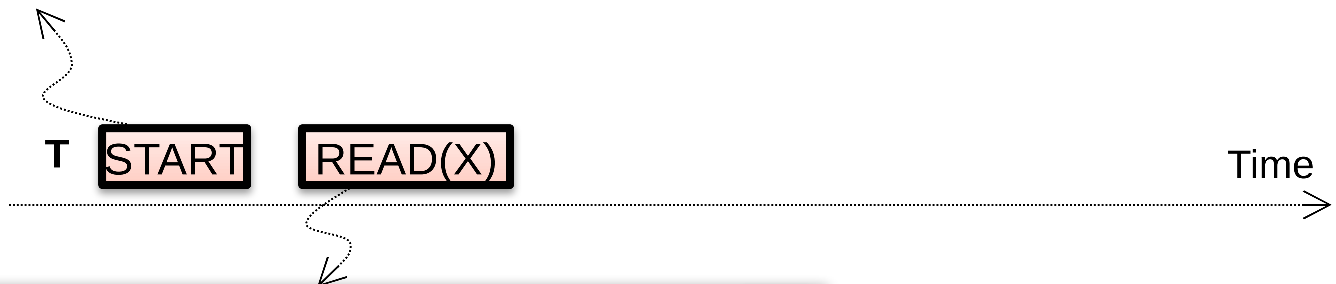
- E.g., we can use OCC for read-write TX, and MVCC for read-only TX
  - Out of the scope of this course

MV-2PL: read-write TX (w/o 2PC & Paxos for simplicity)



MV-2PL: read-only TX

T is assigned a start timestamp T.sts



T reads the biggest version of X(i), such that X(i).cts <= T.sts

Question remains: how do we assign the time to TXs?

Timestamp of start & commit should be assigned in an increasing order

- E.g., global counters with atomic increase
- Note: read-only TX does not necessary increase the time

Example implementation: global counter (on a single machine)

```
u64 global_time; // initial 0
```

```
struct TX {
   u64 start_time;
   u64 commit_time;
   set<...> write_set;
   ...
}
```

```
tx_begin(tx) { // read-only
   tx.global_time =
READ(global_time);
   ...
}

tx_commit(tx) { // read-write
   tx.commit_time =
FAA(global_time);
   ...
```

Question: is global counter suitable for Spanner's use case?

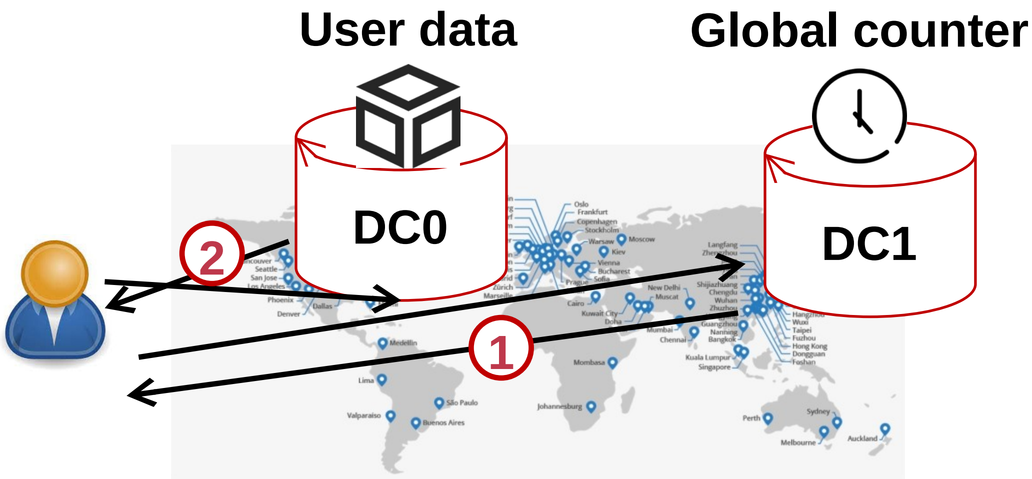
Global time is inefficient for Spanner's use case

Performance overhead

- 1. Extra latency overhead
- 2. Scalability bottleneck

Question: can we avoid reading the global counter?

- Possible when considering the read-only TX



Cache the time locally to avoid querying the global counter

Observation: the read-only TX can read a stale timestamp

- E.g., not reading from the latest global counter
- Still correct
  - TX only reads from a (possible) stale snapshot

Cache global counter to avoid frequently reads

Cache start timestamp for the read-only TX

- The read-only TX no longer needs to do the FAA all the time

Note: here FAA maybe implemented as a remote procedure call

- To the server that stores the global counter

```
u64 global_time; // initial 0
```

```
struct TX {
   u64 start_time;
   u64 commit_time;
   Option<u64> cached_time;
   set<...> write_set;
   ...
}
```

```
tx_begin(tx) {
   if cached_time.is_none():
      tx.cached_time=
Some(READ(global_time));
   tx.start_time =
tx.cached_time.unwrap()
   ...
}
```

Example revisit w/ cached time

```
T1:
Print(A+B)
```

```
T2:
B = 2
A = 3
```

Global counter (initial 2)

T1 (cached_time = 0)

T2

A: A_1 A_0
B: B_1 B_0

3 ← Commit_time = FAA(G)

Write(B, B_3)

Write(A, A_3)

A: A_3 A_1 A_0
B: B_3 B_1 B_0

Start_time = **0** // cached

Read(A) = $\mathbf{A_0}$

Read(B) = $\mathbf{B_0}$

A: A_3 A_1 A_0
B: B_3 B_1 B_0

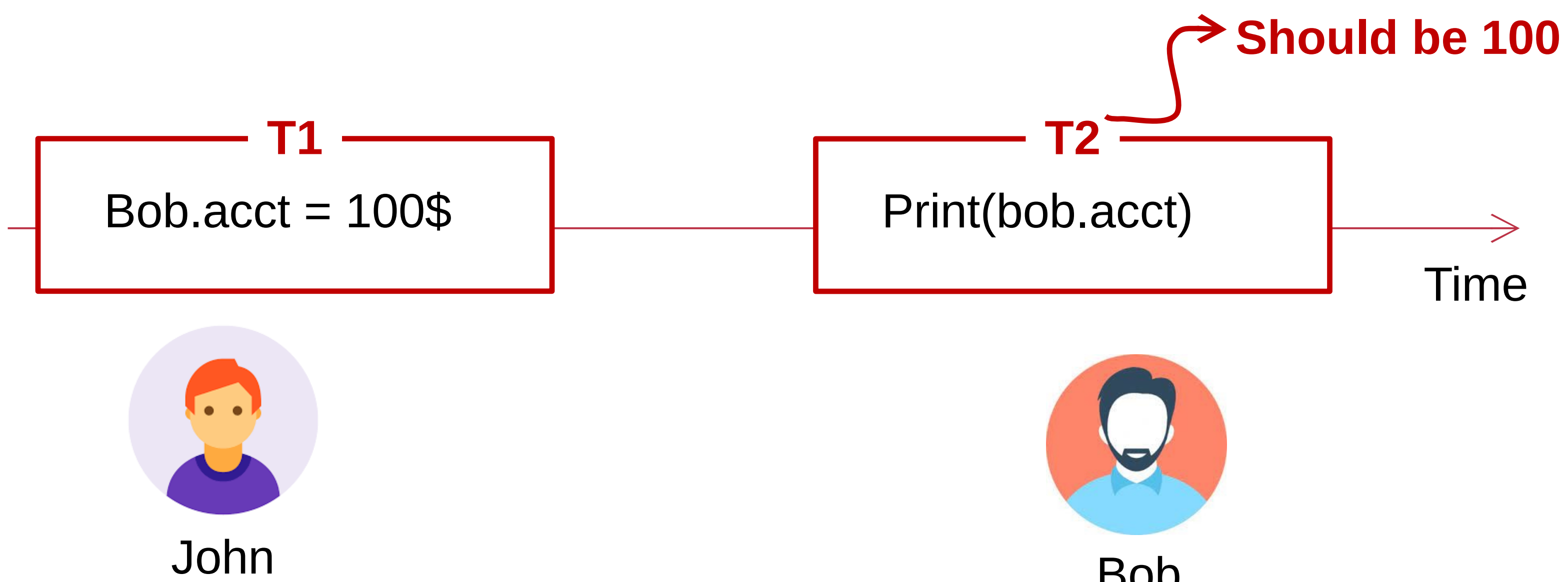
Drawbacks of cached time

No freshness guarantees

– i.e., no external consistency

External consistency is the most desirable for the programmer

– Simplified definition: *If T1 completes before T2 starts, T2 must see T1's writes.*



Drawbacks of cached time

No freshness guarantees

- i.e., no external consistency

Read-write TX still need to acquire the global counter

- Extra latency to communicate with the server that stores the global time
- Possible performance bottleneck

Challenge of the timing in MVCC so far

Time needs synchronization for external consistency (strict serializability)

- If a read-write TX has decided to commit, then a read-only TX starts, then must ensure:
  - $Time_{read\text{-}only} >= Time_{read\text{-}write}$

Global counter trivially satisfies this requirement

- But requires a centralized time server not suitable for geo-replicated databases

Can we use the physical clock of the machines?

- No. Different machines' time are different

Observation: we may not get the accurate physical time, but we can get an accurate **bound**

TrueTime API of Spanner

TrueTime returns a time interval instead of a single point of time

- The interval is a bound. i.e., the time **of the time server** must be in this bound
- The interval is the physical time. i.e., familiar to the user

```
struct timeval {
        time_t tv_sec;
        suseconds_t
tv_usec;
};
int gettimeofday(struct
timeval     *restrict
tv, ...)
```

Linux Time API

```
struct time_interval{
     timeval Lower;
     timeval Upper;
};

int get_truetime(time_interval
*interval);
// Server timeval must be in
[L,U]
```

TrueTime API (Simplified)

Power of TrueTime API (return [L,U])

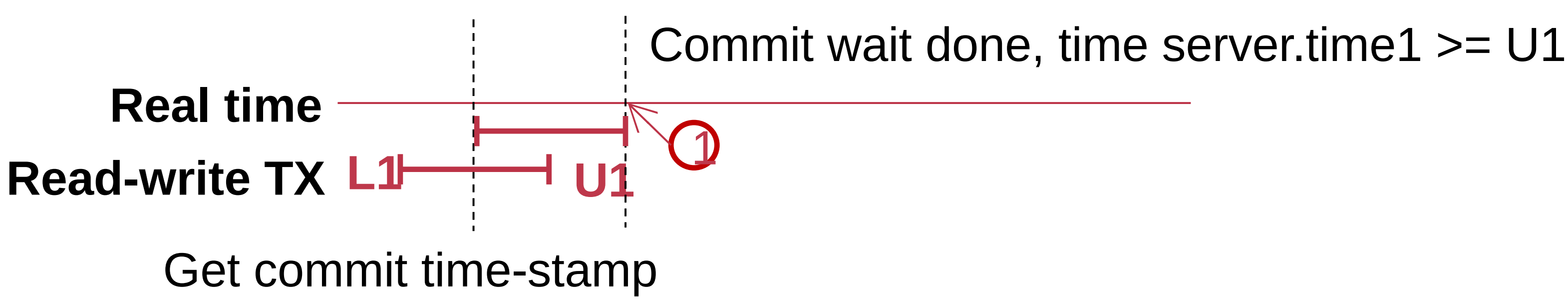
Used to implement external consistency

– If a read-write TX has decided to commit, then:

  - $\text{Time}_{\text{read-only}} >= \text{Time}_{\text{read-write}}$

How to achieve this?

– **Commit wait** for read-write TX: after acquire the commit timestamp, the coordinator wait until (U – L) and uses U as $\text{Time}_{\text{read-write}}$



Power of TrueTime API (return [L,U])

Used to implement external consistency

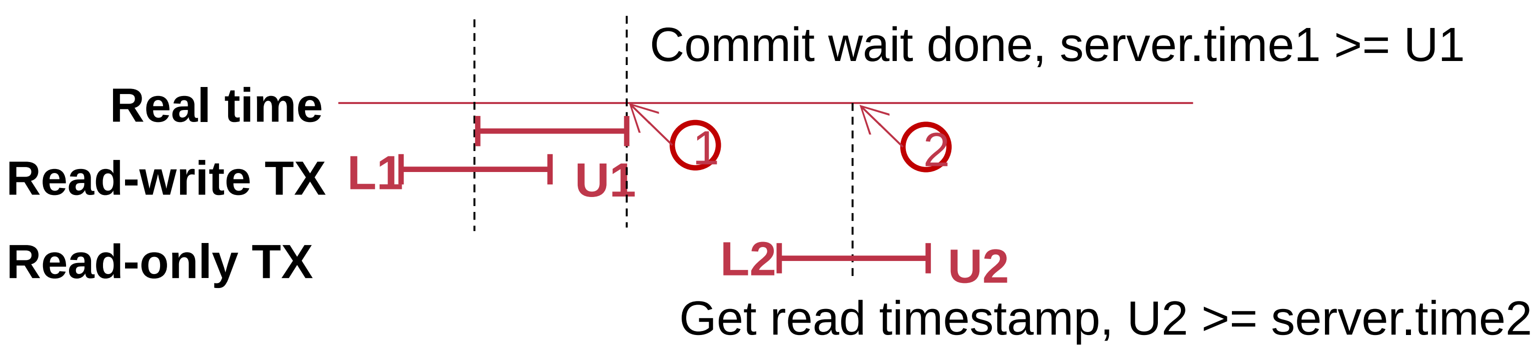
– If a read-write TX has decided to commit, then:

  - $\text{Time}_{\text{read-only}} >= \text{Time}_{\text{read-write}}$

How to achieve this?

– **Commit wait** for read-write TX: after acquire the commit timestamp, the coordinator wait until (U – L) and U1 as $\text{Time}_{\text{read-write}}$

– For read-only TX, simply uses U as the read timestamp



Power of TrueTime API (return [L,U])

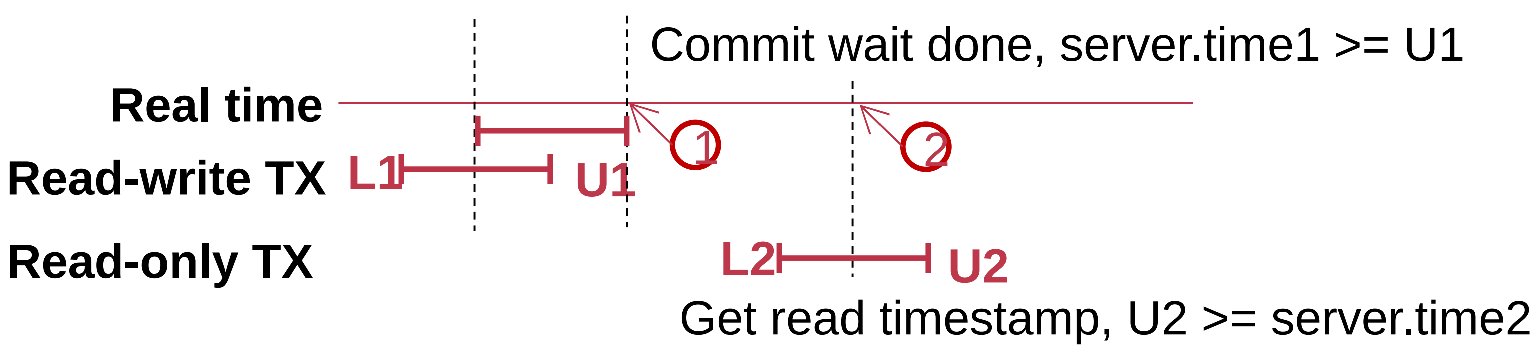
Used to implement external consistency

– If a read-write TX has decided to commit, then:

  - $\text{Time}_{\text{read-only}} >= \text{Time}_{\text{read-write}}$

Correctness

– $\text{Time}_{\text{read-write}}$ (U1) < server.time1 < server.time2 < U2 ($\text{Time}_{\text{read-only}}$)



How to achieve the bound of TrueTime?

TrueTime: how to achieve the bound?

The bound is relative to the time servers as the ground truth

- For simplicity, we first assume there is only one time server

Spanner adopts a variant of Marzullo's algorithm[1]

- Similar to NTP, sync w/ the time server to calculate the bound
- Even single time server will not become the bottleneck: syncing is not on the critical path of the TX's execution

[1] Maintaining the Time in a Distributed System

TrueTime: how to calculate the bound?

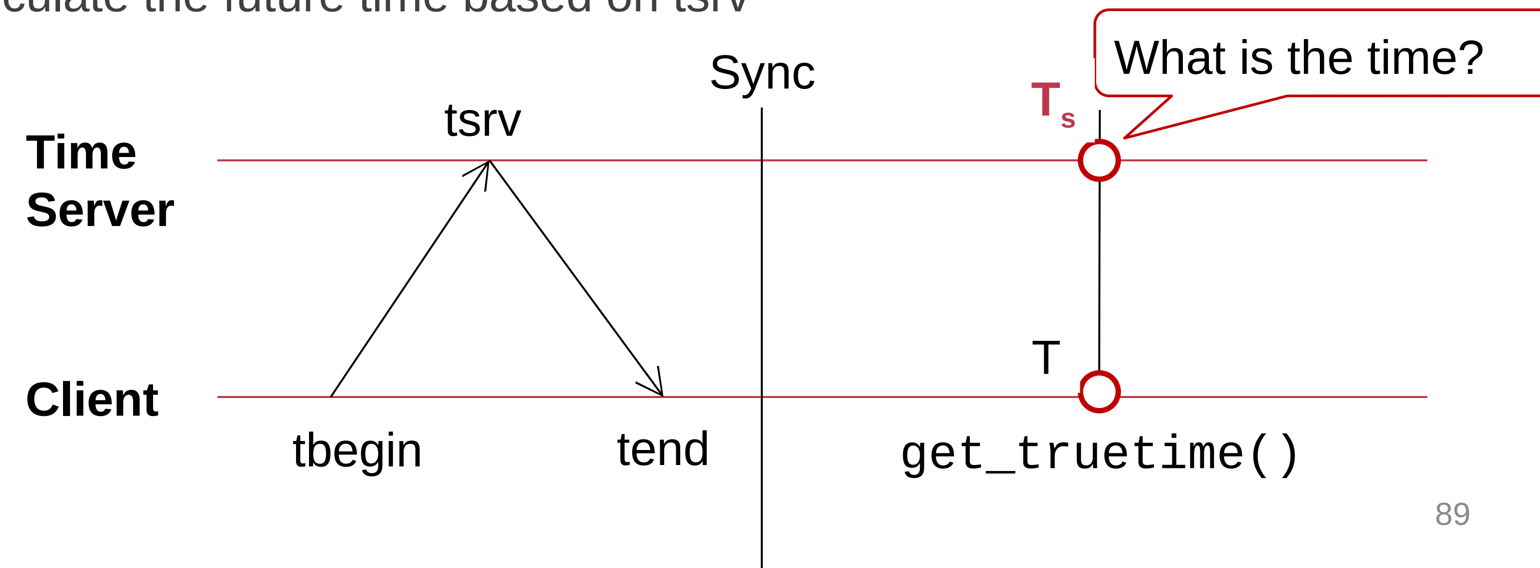
Problem statement

– If the local clock of client is T, what is the time interval of the time server?

High-level idea: send RPC to the server for the query as the measurements!

– But unlike NTP, we don't adjust local time according to tsrv

– We calculate the future time based on tsrv



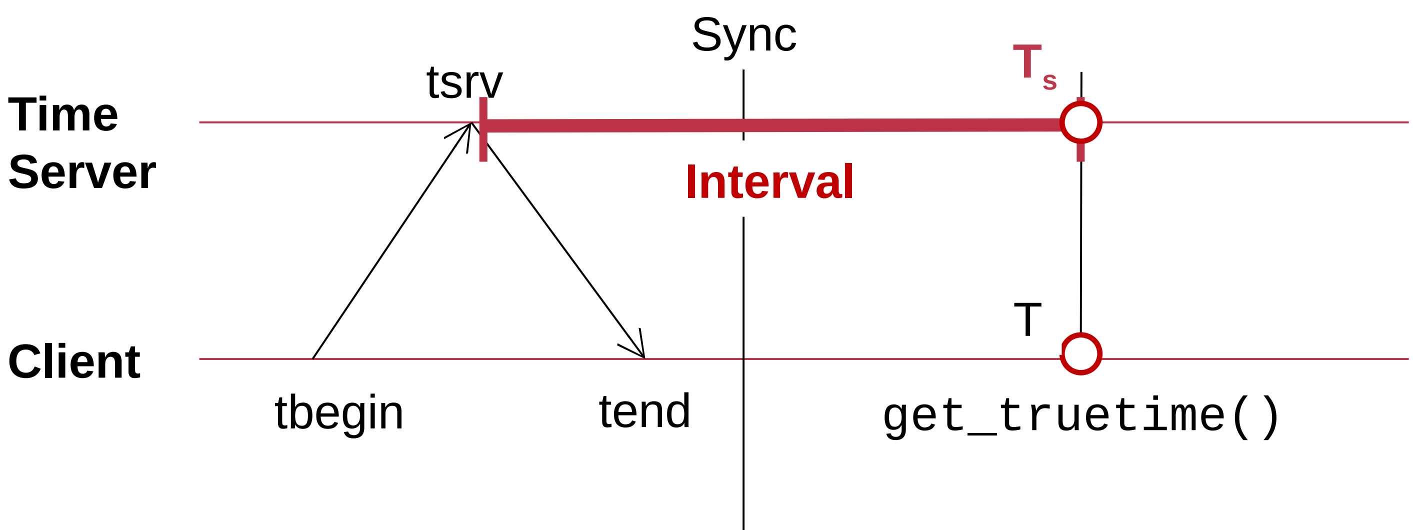
TrueTime: how to calculate the bound?

Simplification: first assuming server advances in the same speed as clients

– Since we know the tsrv, the key question is how to calculate the interval

Question

– How do we calculate the interval based on tsrv, tbegin & tend?



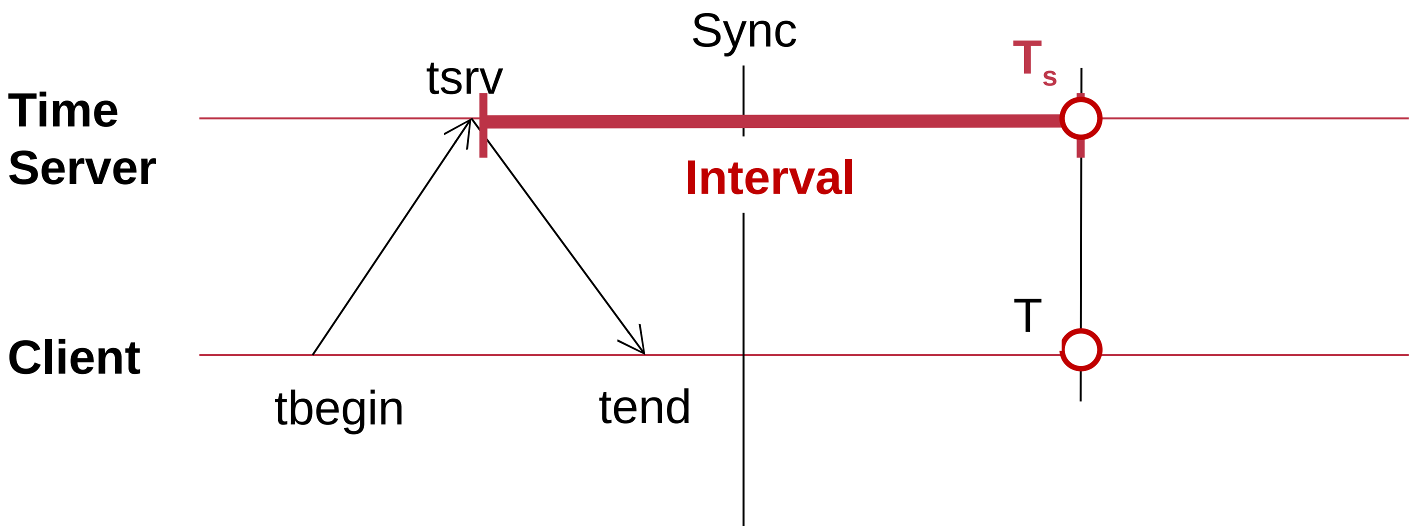
TrueTime: how to calculate the bound?

Since we know the **tsrv**, the key question is how to calculate the **interval**

- Interval >= T – tend
- Interval <= T – tbegin
- T_s = tsrv + Interval

➢ T – tend + tsrv <= T_s <= T – tbegin + tsrv (based on our simplification)

Question: what if there is a drift between client & server?



TrueTime: how to calculate the bound?

~~**Simplification: assuming server advances in the same speed as clients**~~

- $T - tend + tsrv <= T_s <= T - tbegin + tsrv$ (based on our simplification)

Solution: regulate with the drift rate

Assume a fixed ε drift rate between client & server

- After t time, the drift between client & server is **(1 + ε)** or **(1 - ε)**
  - Spanner assumes a fixed as 200us / second

Then the interval is regulated as

- $\boxed{(T - tend) * (1 - \varepsilon) + tsrv} <= T_s <= \boxed{(T - tbegin) * (1 + \varepsilon) + tsrv}$

  L **U**
- Done

Commit wait revisited

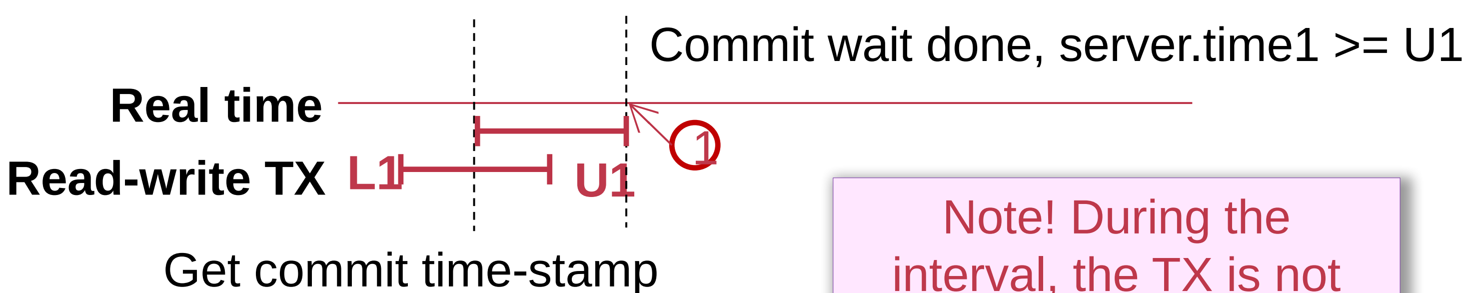
Used to implement external consistency

– If a read-write TX has decided to commit, then:

  - $Time_{read\text{-}only}$ >= $Time_{read\text{-}write}$

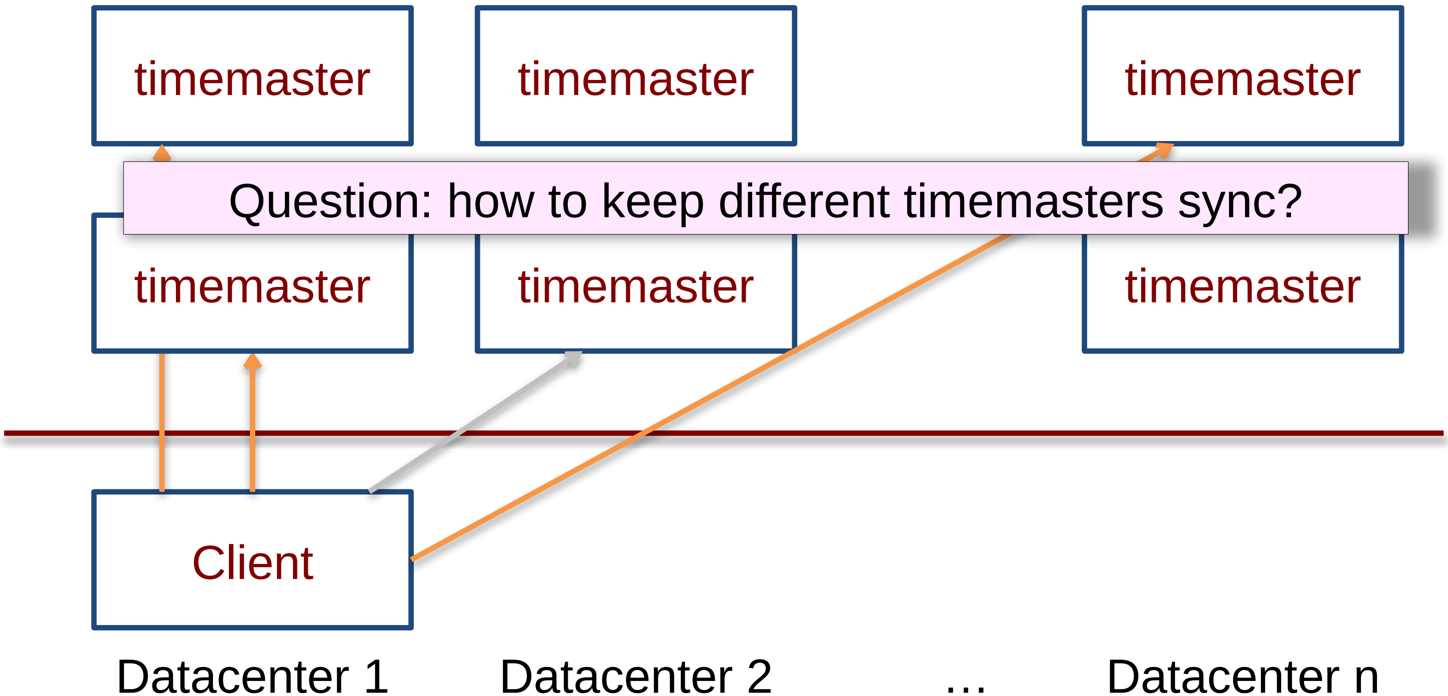
How to achieve this?

– **Commit wait** for read-write TX: after acquire the commit timestamp, the coordinator wait until (U – L)**(1 + ε)** and uses U as $Time_{read\text{-}write}$



Note! During the interval, the TX is not committed

TrueTime adopts multiple time servers

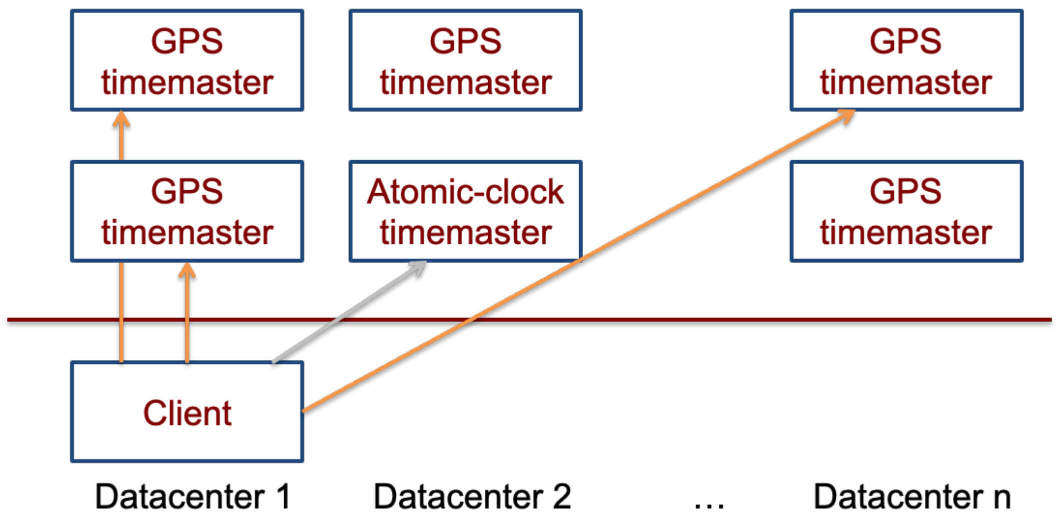


Called timemaster (◀◀) in Spanner

TrueTime adopts multiple time servers

Time servers are backed by GPS & atomic clocks

- High-precision clocks
- E.g., 1 second drift after 20,000,000 year (vs. 200us per second of CPU)
- Atomic clocks are synchronized with each other



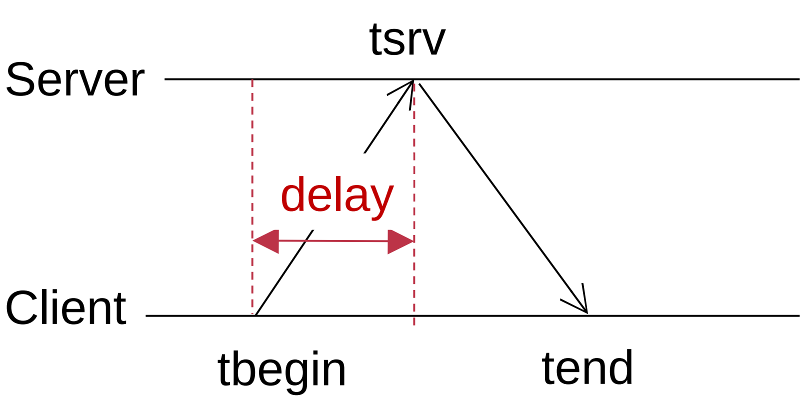
Final takeaway of TrueTime: Network-Induced Uncertainty

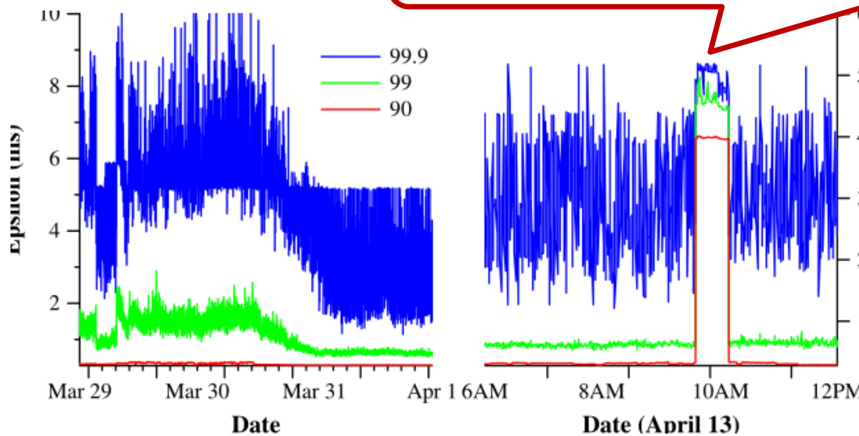
Interval of T

- [(T – tend) * (1 - ε) + tsrv, (T – tbegin) * (1 + ε) + tsrv]
- tsrv – tbegin is rougly estimated as the network delay

Can have spikes if timemasters are out of services

Maintain 2 timemasters





Summary

Reify clock uncertainty in time APIs

– Known unknowns are better than unknown unknowns

– Rethink algorithms (TX’s concurrency control) to make use of uncertainty

Stronger semantics are achievable

– Greater scale != weaker semantics