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Bachelor's Thesis in Informatics

Concurrent Range Locking

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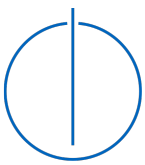

I confirm that this bachelor's thesis in informatics is my own work and I have documented all sources and material used.

München, 15.09.2024 Thua-Duc Nguyen

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Abstract

In modern computing environments, the management of concurrent access to shared resources, such as database tables, file regions, or memory segments, represents a significant challenge. Range locking provides a solution to this challenge by partitioning shared resources into arbitrarily-sized segments, thereby allowing different processes to access these segments concurrently.

Despite its crucial role in various systems, range locking remains an under-researched topic. Existing implementations often suffer from contention and inefficiencies, particularly in highly concurrent environments, underscoring the necessity for more scalable solutions.

In this thesis, we propose a new lock-free concurrent range-locking mechanism to overcome these challenges. Our method improves upon previous designs by eliminating bottlenecks and ensuring high performance in heavily concurrent environments. Our evaluation demonstrates that the proposed method outperforms existing approaches. This thesis provides an in-depth exploration of our method, offering a comprehensive assessment of its effectiveness in modern computing systems.

Contents

List of Figures

List of Tables

1 Introduction

In modern computing environments, managing concurrent access to shared resources, such as database tables, file regions, or memory segments, is a significant challenge. Systems often rely on coarse-grained locking to manage the access between processes. This leads to significant inefficiencies and performance bottlenecks at high concurrency levels. Range locking [1, 2, 3] offers a more refined solution to this challenge. It partitions a shared resource into multiple arbitrarily-sized segments, thereby allowing processes to exclusively acquire and access each segment simultaneously. This strategy effectively addresses the drawbacks and bottlenecks associated with coarse-grained methods, provides more granular control, and improves parallelism.

Range locking in DBMS. In DBMS, range locking ensures data consistency and prevents anomalies such as phantoms, especially at a lower isolation level. They secure not only individual records, but also the gaps between them, preventing other transactions from inserting or modifying records within the range until the transaction is complete. This is particularly challenging in systems where transaction control (TC) and data control (DC) are separated [4]. TC must lock the range before safely interacting with DC, despite not knowing the specific keys involved.

Range locking in file system. In file systems, especially those designed for high-performance computing, are required to handle large-scale parallel I/O requests targeting different parts of a single data file [5, 6]. To achieve this, these systems must apply range locking to efficiently and precisely coordinate concurrent access to shared storage, ensuring consistency and preventing conflicts [1, 7, 2].

Range locking in operating system. Within operating systems, particularly in the Linux kernel community, there has been an increasing interest in applying range locking to alleviate contention issues related to `mmap_sem`. `mmap_sem` is a semaphore used to manage access to virtual memory areas (VMAs) [8, 9, 10]. Because VMAs require synchronized access for operations such as memory mapping, unmapping, and handling page faults, range locking comes into play.

Existing range locking is not efficient. Despite its crucial role in various systems, range locking remains an under-researched topic. Previous implementations of range locking often suffer from contention points due to the reliance on a single lock [11, 3], as a spinlock effectively serializes all incoming lock and unlock requests. Additionally, some methods may be complex and tightly coupled with lock-based concurrency control protocols, which are not applicable for general DBMS operations [12, 13]. These limitations emphasize the need for more refined and scalable solutions to better accommodate the demands of modern, large-scale systems, motivating us to develop an enhanced range lock.

Contribution. In the scope of this thesis, we introduce a new lock-free concurrent range lock. Our method addresses the inefficiencies of previous range locks and ensures high performance in highly concurrent environments. The result shows that, at algorithm level, our method is at least three times faster than existing solutions. This thesis provides an in-depth exploration of the development and evaluation of the proposed range-locking mechanism, offering a thorough understanding of its performance. Furthermore, we compare the proposed solution with state-of-the-art methods, clearly assessing its effectiveness.

2 Related Work

2.1 Coarse-Grained Range Lock

2.1.1 Tree-Based Range Lock

Several works have explored coarse-grained range-locking methods. Jan Kara introduced a range-locking mechanism for the Linux kernel [11], which utilizes a range tree (specifically a red-black tree) to manage range locks and employs a spinlock for synchronization. Each lock is represented as a node in the tree. Similarly, Kim et al. adopted a comparable range lock in their work on pNOVA [7], a variant of the NOVA file system that uses range-based reader-writer locks to enable parallel I/O within a single shared file.

When a thread requests a range lock, it first acquires a spinlock, then traverses the tree to determine the number of locks intersecting with the requested range. Afterward, the thread inserts a node describing its range into the tree and releases the spinlock. If no intersecting locks are found, the thread can proceed with accessing the critical section. If intersecting locks are detected, the thread waits until those locks are released and the number of intersecting locks drops to zero. Upon completing its operation, the thread re-acquires the spinlock, removes its node from the tree, updates the count of overlapping locks, and releases the spinlock. This method ensures that each range is locked only after all previous conflicting range locks have been released, thereby achieving fairness and avoiding livelocks.

Drawbacks

One significant observation is that the coarse-grained spinlock of an interval tree can severely hinder parallelism, as the spinlock effectively serializes all incoming lock and unlock requests. Under heavy concurrent access, this serialization easily becomes a contention point.

Consider three exclusive lock requests for the ranges $A = [1..3]$, $B = [2..7]$, and $C = [4..5]$, arriving in that order. While A holds the lock, B is blocked because it overlaps

with A, and C is blocked behind B. However, in practice, C does not overlap with A and could proceed without waiting. This unnecessary blocking reduces the overall efficiency and concurrency of the system.

2.1.2 List-Based Range Lock

Song et al. [3] introduced a dynamic range-locking design to enhance the implementation of the Linux kernel. Their range lock uses a skip list [14] to dynamically manage the address ranges that are currently locked.

When a thread requests a specific range `[start, start+len)`, the range lock searches the skip list. If an existing or overlapping range is found, it indicates that another thread is currently modifying that range, requiring the requesting thread to wait and then retry. If no overlapping range is found, the requested range is added to the skip list, signifying that the lock has been acquired. Releasing a range involves deleting the corresponding range from the skip list.

Compared to the interval tree, the skip list is more lightweight and efficient, allowing for quicker searches of overlapping ranges.

Drawbacks

Similar to the tree-based range lock, contention remains an issue with this approach. Additionally, it unnecessarily blocks non-conflicting requests, further reducing system efficiency and limiting concurrency.

2.2 Fine-Grained Range Lock

2.2.1 List-Based Range Lock

Kogan et al. [2] introduced a range lock based on a concurrent linked list, where each node represents an acquired range. This design aims to provide a lock-free method, addressing critical shortcomings of previous range-locking implementations. In a lock-free system, processes can proceed without being blocked by locks held by other processes, thereby improving performance and scalability.

The proposed method involves inserting acquired ranges into a linked list sorted by their starting points, ensuring that only one range from a group of overlapping ranges can be inserted using an atomic `compare-and-swap` (CAS) operation. A significant

difference in this method compared to previous ones is that each node has two statuses: `marked` (logically deleted) or `unmarked` (present).

When a thread wants to acquire a range, it iterates through the skip list. If it encounters a marked node, it removes it using CAS and continues to iterate. If the current node protects a range that overlaps, the thread waits until that node is deleted. Otherwise, a node is inserted into the list, signaling that the range is acquired. To release a range, the thread marks the node as deleted.

Drawbacks

Linked List Inefficiency: While this design implements a lock-free method that effectively addresses the limitations of existing range locks, it comes with its own set of trade-offs. In general, insertion and lookup operations in a linked list are less efficient than in tree-like structures. The average time complexity for searching in a linked list is $O(n)$, whereas it is only $O(\log n)$ for skip lists or tree-like structures [15]. Our evaluation will demonstrate that this inefficiency becomes particularly pronounced when handling multiple overlapping ranges within the list.

2.2.2 Bitmap Range Lock

In addition to the tree-based range locking method discussed in Subsection 2.1.1, Kim et al. proposed a lock-free range lock, which they claim offers enhanced efficiency compared to interval tree-based locks. This approach involves dividing a file into segments, each managed by a 32-bit variable that functions as a reader-writer lock. The most significant bit represents the writer lock status (1 for locked, 0 for unlocked), while the remaining 31 bits count the number of active readers. The method utilizes hardware-supported atomic operations to ensure that writer locks can only be set when no other locks are active and that reader locks are granted as long as no writer lock is present. Unlocking is achieved by clearing the writer lock bit and decrementing the reader counter.

Although this method provides finer-grained locking with reduced overhead compared to interval tree-based locks, it is tailored to handle both reader and writer modes and depends on specific memory size constraints. Since our research focuses exclusively on the exclusive mode and does not address the reader-writer combination, we have chosen not to consider this approach as a competitor in our project.

3 Data structure

3.1 Skiplist is suitable for range lock

The effectiveness of a range lock is strongly influenced by the underlying data structure used to manage these ranges. Existing range locks use many different comparison-based data structures, such as interval tree [11, 7] and linked list [3], skip list [2], each of which has advantages and disadvantages.

Interval trees are considered heavyweight compared to linked lists and skip lists because they require complex balancing mechanisms to maintain efficient search times, which adds significant computational and memory overhead. Interval trees also require the maintenance of additional data, such as subtree intervals, making them more resource-intensive. This makes interval trees unsuitable for range locking.

Linked lists have minimal structure and are easy to implement and maintain. However, they have a high cost in terms of Insert/Delete time. The O(n) complexity for the average search case becomes a significant disadvantage when the number of ranges increases,

Skiplists strike a balance between the heaviness of interval trees and the simplicity of linked lists. Despite their theoretically poor worst-case performance, skip lists seldom exhibit such behavior, making them efficient in most scenarios. For example, in a dictionary with more than 250 entries, the probability that a search will take more than three times the expected time is less than one in a million [16].

We strongly believe that the skip list is ideal for implementing range locking.

Data Structure	Best Case	Average Case	Worst Case
Linked List	$O(1)$	$O(n)$	$O(n)$
Skip List	$O(1)$	$O(logn)$	$O(n)$
Interval Tree	$O(\log n)$	$O(\log n)$	$O(\log n)$

Table 3.1: Time complexities of Insert/Delete operations in Linked List, Skip List and Interval Tree

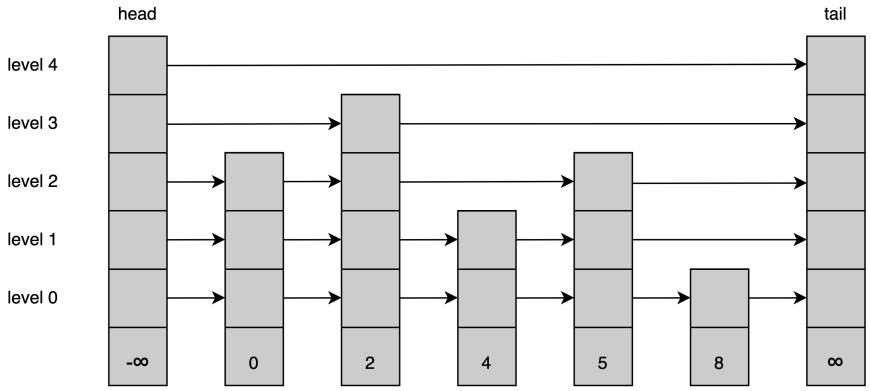


Figure 3.1: An example of a skip list. It has five levels of sorted linked lists.

3.2 Modification of skiplist for concurrent range lock

Although the high-level idea of skiplist meets the requirement for range locking, conventional skiplists does not work out of the box for range locking yet. Firstly, it should be noted that each node of the skiplist stores a key and a value, rather than a range. Secondly, the comparison conditions must be modified, as it is no longer sufficient to simply use the key; instead, both the range start and range end must be taken into account when inserting or deleting. These two aspects – node storage and comparison conditions – require modification.

We therefore propose a new data structure on the basis of the original skip list. We developed our structure based on the `LockFreeSkipList` proposed by Herlihy et al. [17]. The two critical modifications of node storage and comparsion conditions are presented in the Sections 3.2.1 and 4.3.1, respectively.

In summary, `LockFreeSkipList` uses atomic operations (`compareAndSwap()`) to manage node references without locks, thereby enhancing performance in multithreaded environments. When a node is added, the process begins at the lowest level and progresses upward, ensuring immediate visibility. Node removal involves marking nodes from the top down before unlinking them. Additionally, it relaxes the strict structural maintenance of higher levels, focusing on the bottom-level list for set representation, which offers improved scalability and efficiency.

3.2.1 Node

`Node` is the base of our `ConcurrentRangeLock` structure. Each `Node` instance now contains `start` and `end`, which represent one range. `Node` uses an array of `AtomicMarkableReference` (more details in Section 3.2.2) to maintain forward links at each level, which allows for efficient traversal and updates. `Node` provides the following methods:

- `initialize`: sets up a node with specific range and level values.
- `initializeHead`: configures the head node with forward pointers directed to a provided tail node, establishing the initial structure.
- `getTopLevel()`, `getStart()`, `getEnd()`: accessor methods to retrieve the properties of node.

```
1  class Node {
2    private:
3      uint64_t start, end;
4      int topLevel;
5      AtomicMarkableReference** next = nullptr;
6
7    public:
8      void initialize(uint64_t start, uint64_t end, int topLevel) {
9          this->start = start; this->end = end;
10         this->topLevel = topLevel;
11
12         next = new AtomicMarkableReference*[topLevel + 1];
13         for (int i = 0; i <= topLevel; ++i) {
14             next[i] = new AtomicMarkableReference();
15         }
16     }
17
18     void initializeHead(uint64_t start, uint64_t end, int topLevel, Node* tail) {
19         initialize(start, end, topLevel);
20         for (int i = 0; i <= topLevel; ++i){
21             next[i]->store(tail, false);
22         }
23     }
24
25     int getTopLevel() const { return topLevel; }
26     uint64_t getStart() const { return start; }
27     uint64_t getEnd() const { return end; }
28 };
```

Listing 3.1: `ConcurrentRangeLock`: `Node` class

3.2.2 AtomicMarkableReference

We want to keep the high levels of range locking non-blocking. We employed a common technique in concurrent linked list implementations of marking the node [18]. In brief, the least significant bit (LSB) of the next pointer is set to mark the current node as deleted. This enables the soft deletion of a node, which can then be physically deleted at a later stage by any other process without blocking the others. Therefore, we track successors of a `Node` by an array of `AtomicMarkableReference`.

An `AtomicMarkableReference` object encapsulates both a reference to an object of type T and a boolean mark, which can be atomically updated, either together or individually. It uses a single atomic variable, `atomicRefMark`, to store a packed representation of both the reference and a mark. If the mark is `1`, it indicates that the predecessor of this node has been softly deleted. These values are packed and unpacked using bitwise operations, where the least significant bit represents the mark.

Listing 3.2 provides the algorithm for `AtomicMarkableReference`.

The `pack` method combines a node pointer and a boolean mark into a single `uintptr_t` value by encoding the pointer into the lower bits and the mark into the highest bit. Conversely, the `unpack` method decodes this packed value to retrieve the original node pointer and boolean mark.

To atomically set a new node pointer and mark value, the `store` method uses relaxed memory ordering.

The `compareAndSwap` method performs an atomic update of both the reference and mark if they match the expected values, employing acquire-release memory ordering for proper synchronization.

The `attemptMark` method focuses on updating only the mark, provided that the current reference matches the expected one and the mark is different. If the update succeeds, it returns `true`; otherwise, it returns `false`.

The `get` method retrieves the current reference and mark, which stores the mark in the provided boolean pointer. In contrast, the `getReference` method returns the current reference without accessing the mark.

```cpp
class AtomicMarkableReference {
  private:
    std::atomic<uintptr_t> atomicRefMark;

    uintptr_t pack(Node* ref, bool mark) const {
        return reinterpret_cast<uintptr_t>(ref) | (mark ? 1 : 0);
    }

    std::pair<Node*, bool> unpack(uintptr_t packed) const {
        return {reinterpret_cast<Node*>(packed & -1), packed & 1};
    }

  public:
    AtomicMarkableReference() {
        atomicRefMark.store(pack(nullptr, false), std::memory_order_relaxed);
    }

    void store(Node* ref, bool mark) {
        atomicRefMark.store(pack(ref, mark), std::memory_order_relaxed);
    }

    bool compareAndSwap(Node* expectedRef, Node* newRef, bool expectedMark, bool newMark) {
        return atomicRefMark.compare_exchange_strong(
            pack(expectedRef, expectedMark), pack(newRef, newMark), std::memory_order_acq_rel);
    }

    bool attemptMark(Node* expectedRef, bool newMark) {
        auto [currentRef, currentMark] = unpack(atomicRefMark.load(std::memory_order_acquire));
        if (currentRef == expectedRef && currentMark != newMark) {
            return atomicRefMark.compare_exchange_strong(
                current, pack(expectedRef, newMark), std::memory_order_acq_rel);
        }
        return false;
    }

    Node* get(bool* mark) const {
        auto [ref, currentMark] = unpack(atomicRefMark.load(std::memory_order_acquire));
        mark[0] = currentMark;
        return ref;
    }

    Node* getReference() const {
        auto [ref, _] = unpack(atomicRefMark.load(std::memory_order_acquire));
        return ref;
    }
};
```

Listing 3.2: `ConcurrentRangeLock`: `AtomicMarkableReference` class.

4 Approach

In this chapter, we will provide a detailed explanation of the API and algorithm behind the concurrent range lock.

4.1 Concurrent Range Lock API

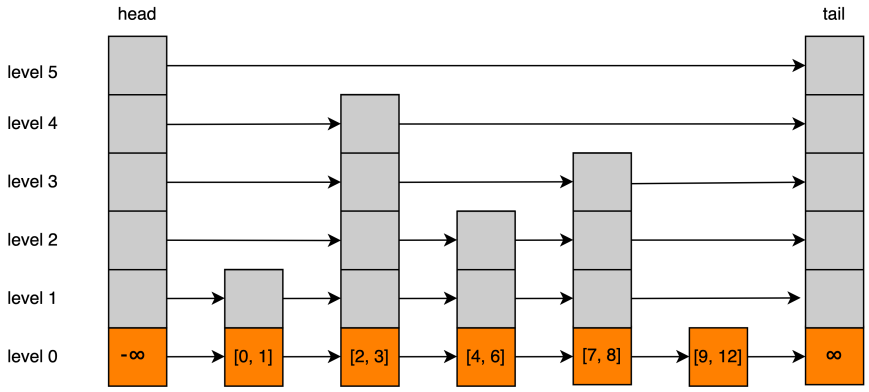


Figure 4.1: `ConcurrentRangeLock`. Each `Node` has a unique range.

The ConcurrentRangeLock class consists of two main functions: `tryLock` and `releaseLock`.

The `tryLock` method attempts to acquire a lock for the specified range `[start, end]`, returning `true` on success and `false` otherwise. The `releaseLock` method releases the lock for the range `[start, end]`, with `true` indicating success and `false` if the range was not found or an error occurred. We will discuss these methods in Subsection 4.2.2 and 4.2.3.

The two primary methods rely heavily on private searching methods such as `findInsert`, `findExact`, and `findDelete`, which handle insertion finding, exact range finding, and physical deletion of ranges, respectively. We will discuss these methods in Section 4.3.1.

```
class ConcurrentRangeLock {
  public:
    ConcurrentRangeLock();

    bool tryLock(uint64_t start, uint64_t end);
    bool releaseLock(uint64_t start, uint64_t end);

  private:
    Node *head, *tail;

    int randomLevel();
    bool findInsert(uint64_t start, uint64_t end, Node **preds, Node **succs);
    bool findExact(uint64_t start, uint64_t end, Node **preds, Node **succs);
    void findDelete(uint64_t start, uint64_t end);
};
```

Listing 4.1: `ConcurrentRangeLock`: API

4.2 Algorithm in details

For the sake of simplicity, we use `uint64_t` as the data strucutre for ranges for our detail algorithm provided in this section. We use the template feature in our open-source C++ implementation to enable generic programming. For further detail, please refer to our open-source code.

4.2.1 Life cycle

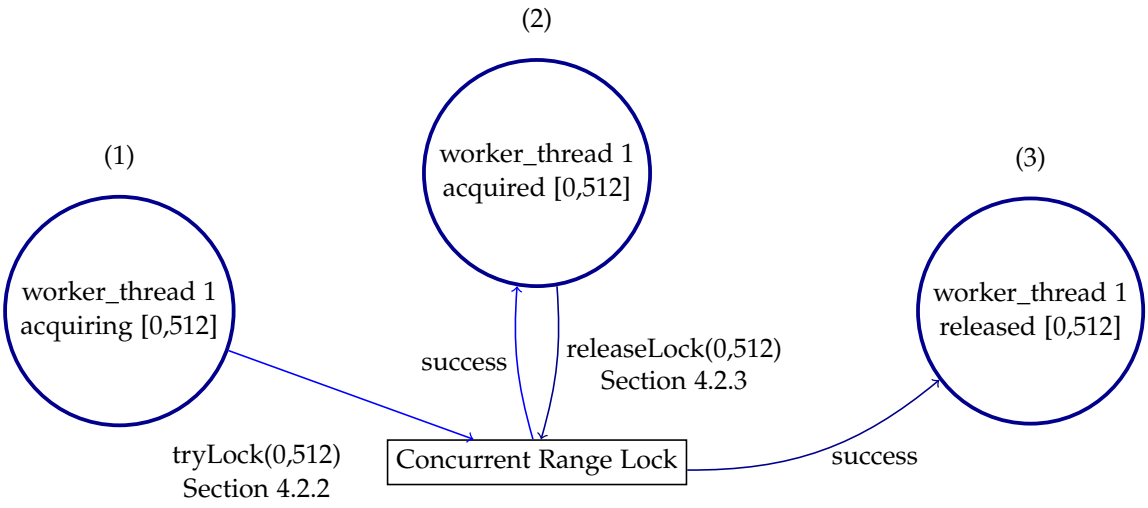


Figure 4.2: `ConcurrentRangeLock`: life cycle one process lock and unlock one range.

Each process has to call the `tryLock` method with the parameters range start, range end, thus the positions on the shared object that this thread wants to acquire. If there is no conflict, that is, if there is no overlap existing on the lock, `tryLock` is successful. Thread now has exclusive access to this segment of the shared object and can freely read and write this segment. After finish using this segment, thread can call `releaseLock` to return his access to this segment. If used correctly, `releaseLock` should always return true.

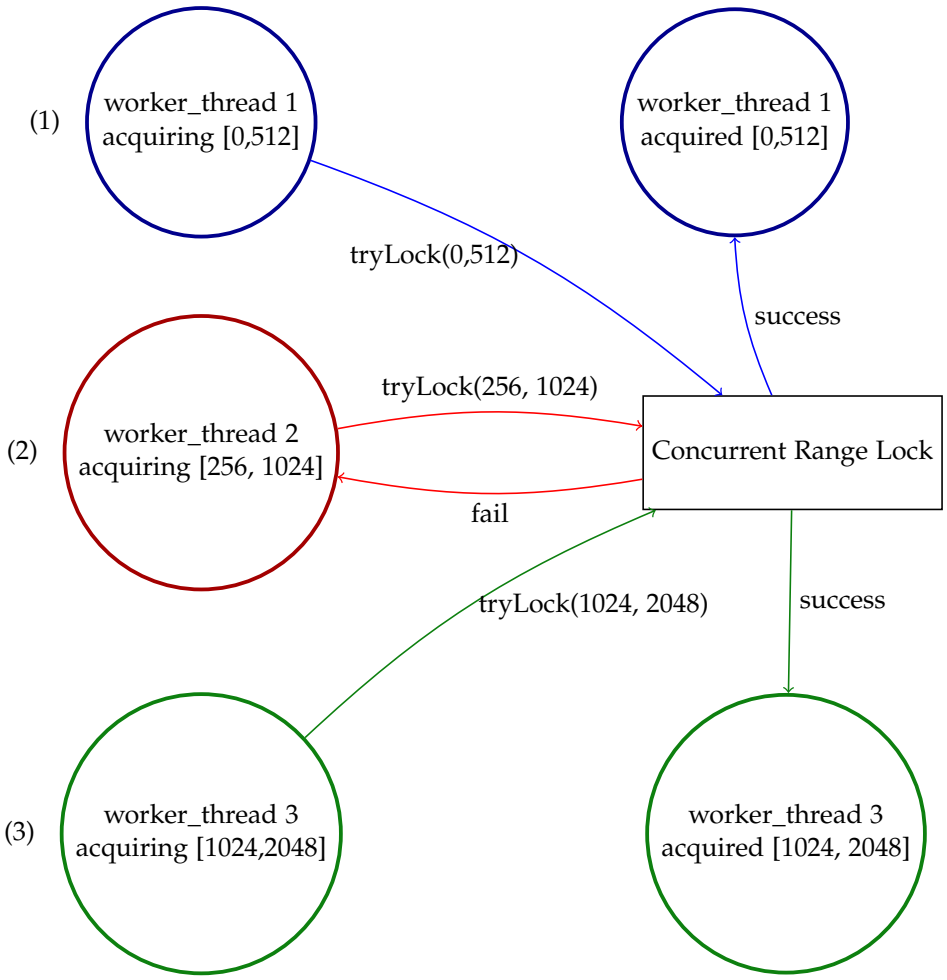


Figure 4.3: The behavior of range lock case: three threads tried to lock ranges. Second thread failed because of range overlap.

4.2.2 Try Lock

The `tryLock` method, shown in Listing 4.2, utilizes `findInsert()` to check if a node with the range `[start, end]` already exists. If found, `tryLock` returns `false`. Otherwise, it creates a new node and attempts to insert it into the list. The node is inserted starting from the bottom level, with `compareAndSwap()` ensuring the integrity of the insertion. If any insertion fails due to concurrent changes, `findInsert()` is called again to update the predecessors and successors, and the process repeats until successful.

```
1   bool ConcurrentRangeLock::tryLock(uint64_t start, uint64_t end) {
2       int topLevel = randomLevel();
3       Node *preds[maxLevel + 1];
4       Node *succs[maxLevel + 1];
5
6       while (true) {
7           if (findInsert(start, end, preds, succs)) {
8               return false;
9           } else {
10              auto newNode = new Node();
11              newNode->initialize(start, end, topLevel);
12
13              for (int level = 0; level <= topLevel; ++level) {
14                  newNode->next[level]->store(succs[level], false);
15              }
16
17              if (!preds[0]->next[0]->compareAndSwap(succs[0], newNode, false, false)) {
18                  continue;
19              }
20
21              for (int level = 1; level <= topLevel; ++level) {
22                  while (!preds[level]->next[level]->compareAndSwap(
23                      succs[level], newNode, false, false)) {
24                      findInsert(start, end, preds, succs);
25                  }
26              }
27
28              return true;
29          }
30      }
31  }
```

Listing 4.2: `ConcurrentRangeLock: tryLock()`.

Algorithm in details

The `tryLock` method, shown in Listing 4.2, uses `findInsert()` to determine whether a node with range `[start, end]` is already in the list (line 7). `tryLock` also calls `findInsert()` to initialize the preds[] and succs[] arrays to hold the ostensible predeces-

sors and successors of the new node. If an unmarked node is found in the bottom-level list, `findInsert()` returns `true` and the `tryLock` method returns `false`, indicating that part of the target range is already acquired by other thread. See Figure 4.3 as an example. The unsuccessful linearization point of `tryLock` is the same as the successful linearization point of `findInsert()` (line 8). If no unmarked node is found, the next step is to add a new node with the target range into the structure.

A new node is created with a randomly chosen top-level (lines 10–11). The next references of the new node are unmarked and set to the successors returned by the `findInsert()` method (lines 13–15). The next step is to try to add the new node by linking it into the bottom-level list between the `preds[0]` and `succs[0]` nodes returned by `findInsert()`. We use the `compareAndSwap()` method to set the reference while validating that these nodes still refer one to the other and have not been removed from the list (line 17). If the `compareAndSwap()` fails, something has changed and the call restarts. If the `compareAndSwap()` succeeds, the item is added, and line 17 is the call's linearization point. The findInsert() then links the node in at higher levels (lines 21–25). For each level, it attempts to splice the node by setting the predecessor, if it refers to the valid successor, to the new node (lines 22–23). If successful, it breaks and moves on to the next level. If unsuccessful, then the node referenced by the predecessor must have changed, and `findInsert()` is called again to find a new valid set of predecessors and successors (line 24). We discard the result of calling `findInsert()` because we care only about recomputing the ostensible predecessors and successors on the remaining unlinked levels. The method returns true once all levels are linked (line 27).

4.2.3 Release Lock

The `releaseLock` method first calls `findExact()` to locate an unmarked node with the specified range `[start, end]`. If found, it marks all node levels except the bottom one, preparing the node for removal. The method then attempts to mark the bottom-level link using `compareAndSwap()`, which serves as the linearization point for a successful `releaseLock`. If marking fails due to concurrent modifications, the method retries or returns false, depending on the state of the node.

Algorithm in details

The `releaseLock` method, shown in Listing 4.3, calls `findExact()` to determine whether an unmarked node with a matching range `[start, end]` is in the bottom-level list (line 7). If no node is found in the bottom-level list, or if the node with a matching

```
1  bool ConcurrentRangeLock::releaseLock(uint64_t start, uint64_t end) {
2      Node *preds[maxLevel + 1];
3      Node *succs[maxLevel + 1];
4      Node *succ;
5
6      while (true) {
7          bool found = findExact(start, end, preds, succs);
8          if (!found) {
9              return false;
10         } else {
11             Node *nodeToRemove = succs[0];
12             for (int level = nodeToRemove->getTopLevel();
13                     level >= 0 + 1; level--) {
14                 bool marked[1] = {false};
15                 succ = nodeToRemove->next[level]->get(marked);
16                 while (!marked[0]) {
17                     nodeToRemove->next[level]->attemptMark(succ, true);
18                     succ = nodeToRemove->next[level]->get(marked);
19                 }
20             }
21
22             bool marked[1] = {false};
23             succ = nodeToRemove->next[0]->get(marked);
24             while (true) {
25                 bool iMarkedIt = nodeToRemove->next[0]->compareAndSwap(
26                         succ, succ, false, true);
27                 succ = succs[0]->next[0]->get(marked);
28                 if (iMarkedIt) {
29                     findDelete(start, end);
30
31                     return true;
32                 } else if (marked[0]) {
33                     return false;
34                 }
35             }
36         }
37     }
38 }
```

Listing 4.3: `ConcurrentRangeLock: releaseLock()`

range `[start, end]` is marked, the method returns false. The linearization point of the unsuccessful `releaseLock` is that of the `findExact()` method called in line 7.

If an unmarked node is found, the method logically removes the associated key from the abstract set and prepares it for physical removal. This step uses the set of ostensible predecessors (stored by `findExact()` in `preds[]`) and the victim (returned from `findExact()` in `succs[]`). First, starting from the top-level, all links up to and **not including** the bottom-level link are marked (lines 12–20) by repeatedly reading next and its mark and applying `attemptMark()`. If the link is found to be marked (either because it was already marked or because the attempt succeeded), the method moves on to the

next-level link. Otherwise, the current level's link is reread since another concurrent thread must have changed it, so the marking attempt must be repeated.

Once all levels but the bottom one have been marked, the method marks the bottom level's next reference. If successful, this marking (line 27) is the linearization point of a successful `releaseLock`. The `releaseLock` method tries to mark the next field using `compareAndSwap()`. If successful, it can determine that it was the thread that changed the mark from false to true. Before returning true, the `findDelete()` method is called. This call is an optimization: `findDelete()` physically removes all links to the node it is searching for.

On the other hand, if the `compareAndSwap()` call fails, but the next reference is marked, then another thread must have concurrently removed it, so `releaseLock` returns false. The linearization point of this unsuccessful `releaseLock` is the linearization point of the `releaseLock` method by the thread that successfully marked the next field. Notice that this linearization point must occur during the `releaseLock` call because the `findExact()` call found the node unmarked before it found it marked.

Finally, if the `compareAndSwap()` fails and the node is unmarked, the next node must have changed concurrently. Since the victim is known, there is no need to call find() again, and `releaseLock` simply uses the new value read from next to retry the marking.

4.3 Utility

4.3.1 Find

Both `tryLock` and `releaseLock` methods rely heavily on `find` methods. There are several find methods in our implementation that serve different purposes:

- `bool findInsert(uint64_t start, uint64_t end, Node** preds, Node** succs)`: checks if the target range `[start, end]` is free to be inserted.
- `bool findExact(uint64_t start, uint64_t end, Node** preds, Node** succs)`: checks if the target range `[start, end]` is already present in the skip list.
- `void findDelete(uint64_t start, uint64_t end)`: finds the target range `[start, end]` from the skip list to physically delete the node which contains the corresponding range.

These `findInsert` and `findExact` methods also fill in the `preds[]` and `succs[]` arrays with the target node's predecessors and successors at each level. Because the goal of `findDelete` is only to snip out all the deleted node, there is no need to fill any array.

Nevertheless, these methods have to maintain the following two properties:

- During traversal, they need to skip over marked nodes. They use `compareAndSwap()` to ensure that they remove all softly deleted node on the way.
- Every `preds[]` reference is to a node with a key strictly less than the target.

Algorithm in details

The `find()` method starts by traversing the LockFreeSkipList from the `topLevel` of the head sentinel, which has the maximal allowed node level. It proceeds down the list level by level, filling in the `preds` and `succs` nodes. These nodes are repeatedly advanced until `pred` refers to a node with the `end` value on that level that is strictly less than the target range `end` (lines 13–29).

While traversing, it repeatedly snips out marked nodes from the current level as they are encountered (lines 15–22) using a `compareAndSwap()`. `compareAndSwap()` function also validates that the next field of the predecessor still references the current node.

Once an unmarked `curr` node is found (line 23), it is tested to see if its `start` is greater than or equal to the target `start`. If so, `pred` is advanced to `curr`, `curr` is advanced to

```
1  bool ConcurrentRangeLock::find(uint64_t start, uint64_t end,
2      Node **preds, Node **succs) {
3      bool marked[1] = {false};
4      bool snip;
5      Node *pred, *succ, *curr;
6
7    retry:
8      while (true) {
9      pred = head;
10     for (int level = maxLevel; level >= 0; level--) {
11         curr = pred->next[level]->getReference();
12
13         while (start > curr->getStart()) {
14             succ = curr->next[level]->get(marked);
15             while (marked[0]) {
16                 snip = pred->next[level]->compareAndSwap(curr, succ, false, false);
17
18                 if (!snip) goto retry;
19
20                 curr = pred->next[level]->getReference();
21                 succ = curr->next[level]->get(marked);
22             }
23             if (start >= curr->getStart()) {
24                 pred = curr;
25                 curr = succ;
26             } else {
27                 break;
28             }
29         }
30
31         preds[level] = pred;
32         succs[level] = curr;
33     }
34
35     return **condition**;
36     }
37 }
```

Listing 4.4: `ConcurrentRangeLock`: `find` methods.

`succ`, and the traverse continues. Otherwise, the current range of `pred` is the immediate predecessor of the target node. The `find()` method then breaks out of the current level search loop, saving the current values of `pred` and `curr` (lines 26–32).

The `find()` method continues this process until it reaches the bottom level. An important point is that each level's traversal maintains the previously described properties. Specifically, if a node with the target key is in the list, it will be found at the bottom level even if nodes are removed at higher levels. When traversal stops, `pred` refers to a predecessor of the target node. The method descends to each next lower level without skipping over the target node. If the node is in the list, it will be found at the bottom level. Additionally, if the node is found, it cannot be marked because if it were marked,

it would have been snipped out in lines 15–22. Thus, the condition test on line 35 only needs to check if there are overlap ranges (`findInsert`) or if the start and end of the node match the target start and end (`findExact`).

1. `findInsert`:

```
return (!(start > pred->getEnd() && end < curr->getStart()));
```

2. `findExact`:

```
return (start == curr->getStart() && end == curr->getEnd());
```

The linearization points for both successful and unsuccessful calls to the `find()` method occur when the `curr` reference at the bottom-level list is set, either at line 11 or line 20, for the last time before the success or failure of the `find()` call is determined at line 35.

5 Evaluation

This section will evaluate our proposed concurrent range lock under different scenarios. The goal is to see the scalability and throughput of our mechanism compared to state-of-the-art range locks.

5.1 Competitors

We denote our primary implementation as `Our`. In addition to `Our`, we implement two different competitors for comparison. The first competitor is a scalable range lock proposed by Kogan et al. [2], denoted as `Scalable RL`. We specifically implemented the Exclusive Access variant presented in their paper, as it aligns with the focus of our research. The second competitor is a skip list range lock proposed by Song et al. [3], which we denote as `Song RL`.

5.2 Benchmark enviroment

For benchmark purposes, we utilized a server with an AMD Ryzen 9 7950X processor, which features 16 cores and 32 threads, providing computational power for the experiments. The server configuration included a virtual memory cache with a capacity of 128 GB, backed by 32 GB of physical memory.

5.3 Microbenchmark

5.3.1 Workload

The primary objective of a concurrent range lock mechanism is to enable multiple threads to access disjoint parts of the same shared object efficiently. To simulate and evaluate the effectiveness of different range locking strategies, we utilize the `mmap()` system call to create a shared object in memory and the `memset()` function to simulate write operations to this shared object. For each range, we will write 1KB. The use of

`memset()` serves as a placeholder for actual modifications, enabling us to focus on the performance characteristics of the locking mechanism itself.

To explore various levels of contention and potential usage scenarios for range locks, we have devised two distinct workloads, each designed to stress the locking mechanism under different conditions:

- `W1`**:** In this workload, each thread operates with fine granularity. A thread locks a single memory range, performs a modification (simulated by `memset()`), and immediately releases the lock before proceeding to the next range in its queue. This approach simulates a scenario with minimal contention, focusing on the efficiency of the lock mechanism in handling rapid lock acquisition and release cycles. The primary objective is to assess the overhead introduced by the locking mechanism under low contention and to evaluate its performance in scenarios demanding high throughput with minimal waiting times.
- `W2`**:** This workload introduces a more complex and realistic scenario. Here, threads perform batched memory operations, where a series of memory ranges (typically 16) are locked, modified (simulated by `memset()`) and then unlocked in a single batch. The goal is to test all competitors under heavier threaded conditions. By locking and unlocking in batches, the number of ranges within the data structure increases, providing insight into how efficiently each competitor can search, lock, and unlock ranges as the load increases. This workload emphasizes the performance of the locking mechanism in handling real-world use cases where contention may be higher.

Through these workloads, we aim to comprehensively evaluate the performance of the concurrent range lock mechanism under different workloads and contention scenarios. Each workload provides insights into the lock's scalability, efficiency, and overall robustness in handling varying degrees of parallelism and contention.

5.3.2 Optimal Height for Range Locking

The skip list height in our range-locking mechanism plays a critical role in balancing performance and resource utilization. A skip list with insufficient height may need to optimize lookup times effectively, leading to slower performance. In contrast, an excessive height increases memory consumption and management overhead without proportionate gains in efficiency. This test aims to determine the optimal height for our concurrent range-locking mechanism before comparing it to alternative approaches. We conducted experiments using the two workloads described in Subsection 5.3.1 to identify the best skip list height.

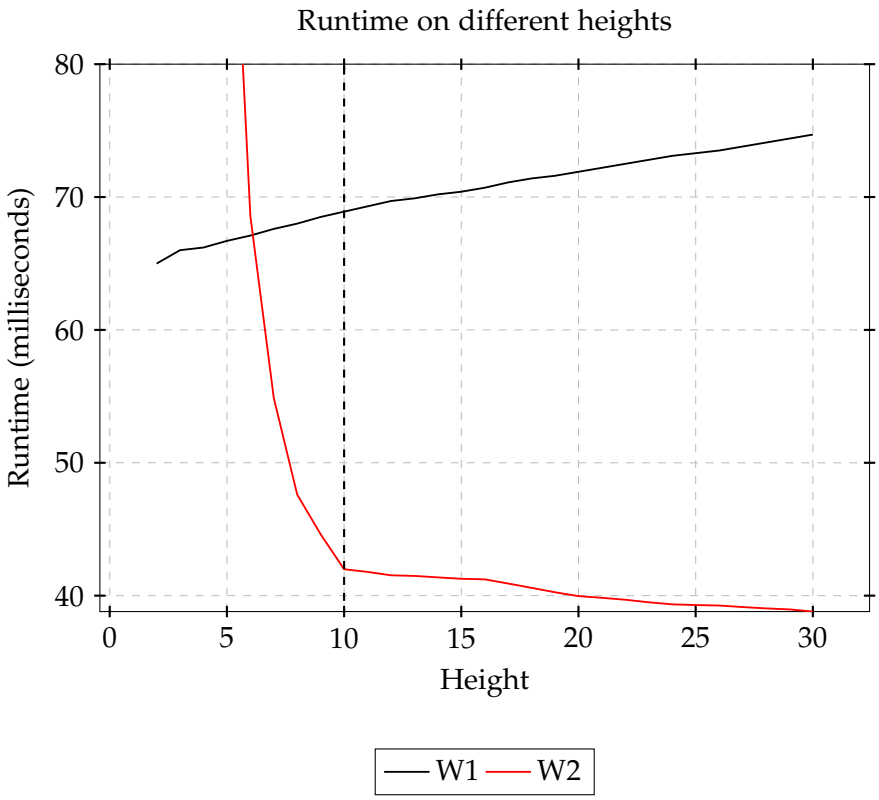


Figure 5.1: Runtime for different heights of range lock

For **W1**, where contention is minimal, and each thread locks and releases ranges immediately, the height of the skip list has negligible impact on performance. In this scenario, since the number of ranges locked at any given time is limited to the number of active worker threads, the benefits of a taller skip list are not realized. The overhead of managing additional levels outweighs any potential gains, leading to similar performance across different heights.

For **W2**, the total runtime decreases as the skip list height increases, aligning with the expected behavior. The multi-level structure of skip lists allows for faster lookups, reducing runtime as height increases. However, beyond a height of 10, further increases yield diminishing returns. This is due to the probabilistic level assignment in skip lists, where the chance of a node reaching higher levels decreases exponentially (e.g., level 12 has a probability of about $\frac{1}{2^{12}} \approx 0.00024\%$). The sparsely populated upper levels contribute little to performance improvement but increase memory usage.

Based on these observations, we identify a sweet spot at a height of `10`, where the trade-

off between performance and resource utilization is optimized. This height balances the need for efficient lookups with manageable memory overhead, making it the ideal choice for our range-locking mechanism in subsequent benchmarks.

5.3.3 Result

Figure 5.2 shows the number of successful locks and unlocks per second with increasing working threads under **W1**. We can see that `Our` has good scalability and outperforms the other two. We archive four times more operations than `Scalable RL` and twelve times more than `Song RL`. The poor performance of `Song RL` is due to the immediate locking and releasing mechanism, combined with the low memory overhead of `memset` (1KB per range). The spinlock in `Song RL` becomes a significant point of contention, leading to its poor performance. Additionally, since the locks are released immediately, `Song RL` cannot effectively leverage its skip list data structure, as the maximum number of ranges in the list is limited to the number of worker threads. However, in more realistic scenarios, such as those discussed in Section 5.4 or **W2**, we will see `Song RL` perform much better.

Figure 5.3 illustrates the same result for **W2**, `Our` continued to outperform the others, achieving two to nine times more operations than `Scalable RL` and two to six times more than `Song RL`. In this workload, the number of ranges locked simultaneously increases to the number of threads multiplied by the batch size (16). This allowed `Song RL` to perform significantly better compared to `Scalable RL`. However, `Scalable RL` demonstrated better scalability as the number of worker threads increased. Beyond sixteen threads, `Scalable RL` began to surpass `Song RL` in performance.

In summary, `Our` consistently outperforms both `Scalable RL` and `Song RL` across different workloads. While `Song RL` struggles with contention in **W1**, it performs better in more complex scenarios like **W2**. However, `Scalable RL` demonstrates better scalability with increasing thread counts, eventually surpassing `Song RL` as threads increase.

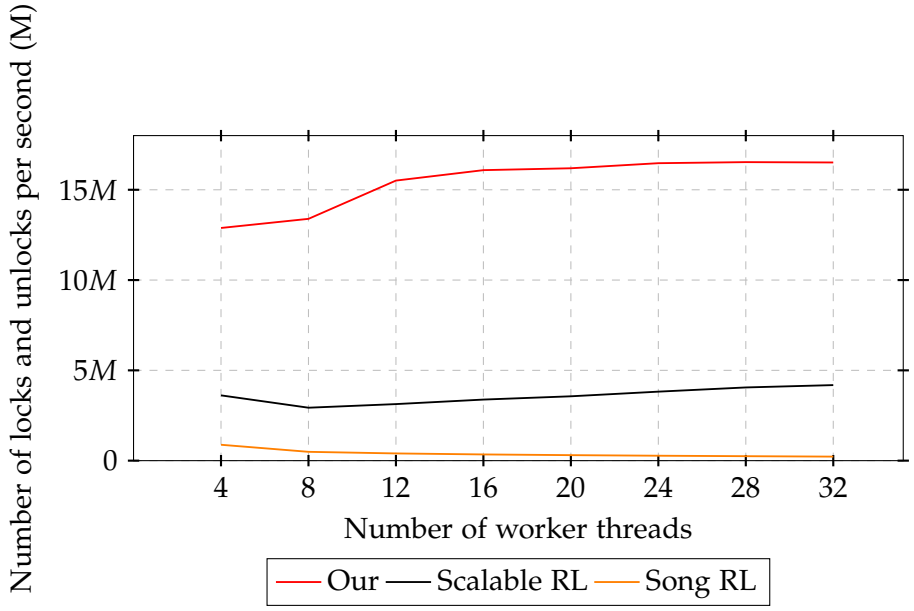


Figure 5.2: **Workload 1**: Lock and unlock operations per second. `Our` achieves 4x more operations than `Scalable RL` and 12x more than `Song RL`.

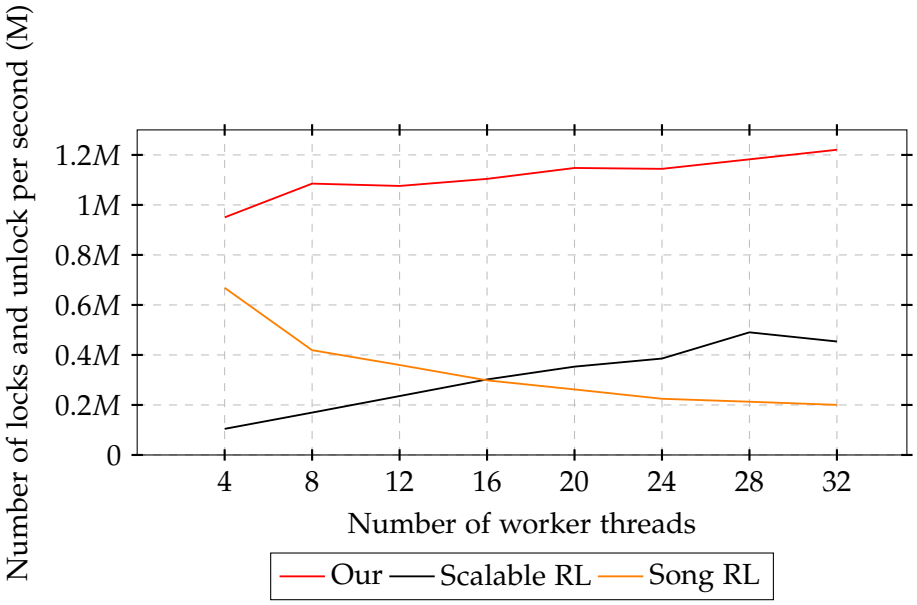


Figure 5.3: **Workload 2**: Lock and unlock operations per second. `Our` achieves 2-9x more than `Scalable RL` and 2-6x more than `Song RL`.

5.4 Evaluation with real use-cases

While microbenchmarks are useful for evaluating and comparing range locks at algorithm level, they often do not capture the complexities of real-world workloads. Testing in real system allows us to move beyond synthetic benchmarks and evaluate the effectiveness of range locks in a live, operational environment where factors like transaction throughput, latency, and contention can significantly impact performance. This ensures that the range locks are not only theoretically efficient but also practically robust under realistic conditions. For those reason, we run an other benchmark of range lock integrated into `Leanstore`.

5.4.1 Leanstore

`Leanstore` [19] is a high-performance storage engine designed to support various database management systems. To further enhance its capabilities, Nguyen et al. [20] introduced a comprehensive design for allocating and logging large objects, which has been integrated into Leanstore. Their performance study demonstrates that this approach not only outperforms many popular file systems but also ensures transactional consistency and durability for large objects. Given the crucial role that range locks play in their design, we integrated our concurrent range lock mechanism to enable realistic and rigorous benchmarking.

In the proposed design, range locks synchronize access to shared aliasing areas, which are contiguous ranges of virtual memory addresses used to present disjointed extents as contiguous memory. When a worker needs to allocate virtual memory for large BLOBs, particularly when these BLOBs exceed the size of the worker-local aliasing area, it must reserve free virtual memory from the shared aliasing area. A range lock mechanism prevents concurrent workers from accessing overlapping memory regions. The range lock operates by locking specific ranges within the shared aliasing area, ensuring that only one worker can modify or access a particular memory range. This prevents race conditions and ensures data consistency while simultaneously allowing multiple workers to operate on different memory ranges.

5.4.2 Competitors

We integrated all three versions described in Section 5.3. Additionally, we included the original, specifically optimized range lock version by the author as a fourth competitor. We denoted it as `Bitmap`

`Bitmap` manages locking over a range of blocks in a buffer. It iterates through the specified block range in chunks, calculating which bits in the lock need to be updated for each chunk. It tries to set the corresponding bits case lock and clear bits case unlock in an atomic manner. If a bit was already set by another thread, the function detects this conflict, aborts the current operation, and undoes any changes made so far.

`Bitmap` is lightweight and efficient implementation of range locking. However, it does not support locking and unlocking a specific range from page x to page y, but rather an entire block of 1GB. Since our research focuses exclusively on the generalized use case of range locking, we have chosen not to consider this approach as a competitor in our project.

5.4.3 Workload

The experiment, utilizing synthetic YCSB [21] workloads, was designed to evaluate the performance of LeanStore under a read-intensive, multithreaded environment. The workload comprises exclusively read operations, executed over a 10-second duration, with each read operation employing a straightforward `memcpy()` function. Each record in the workload has a payload size of 1 MB, with 1000 records being processed. The experiment also incorporates a buffer management configuration that allocates 128 GB of virtual and 32 GB of physical memory, providing a robust test of LeanStore's ability to manage large data sets under constrained physical memory conditions.

5.4.4 Result

Throughput

We evaluated the throughput of four competitors as the number of worker threads increased. The result is depicted in Figure 5.4. As anticipated, the `Bitmap` implementation consistently outperformed the other three competitors. This superiority can be attributed mainly to its optimized use of the aliasing arena, making it the most efficient workload handling. Our implementation (`Our`) demonstrated competitive performance, closely trailing `Bitmap` in terms of transactions per second. Both `Bitmap` and `Our` showed effective scalability with an increasing number of worker threads, though `Bitmap` maintained a consistent, albeit slight, performance advantage.

Conversely, `Song RL` initially exhibited substantial throughput but encountered a sharp decline beyond 24 worker threads. This decline underscores the limitations of its coarse-grained range lock mechanism, as discussed in Subsection 2.1.2. The single spinlock in `Song RL` becomes a significant bottleneck in high-contention scenarios, leading to

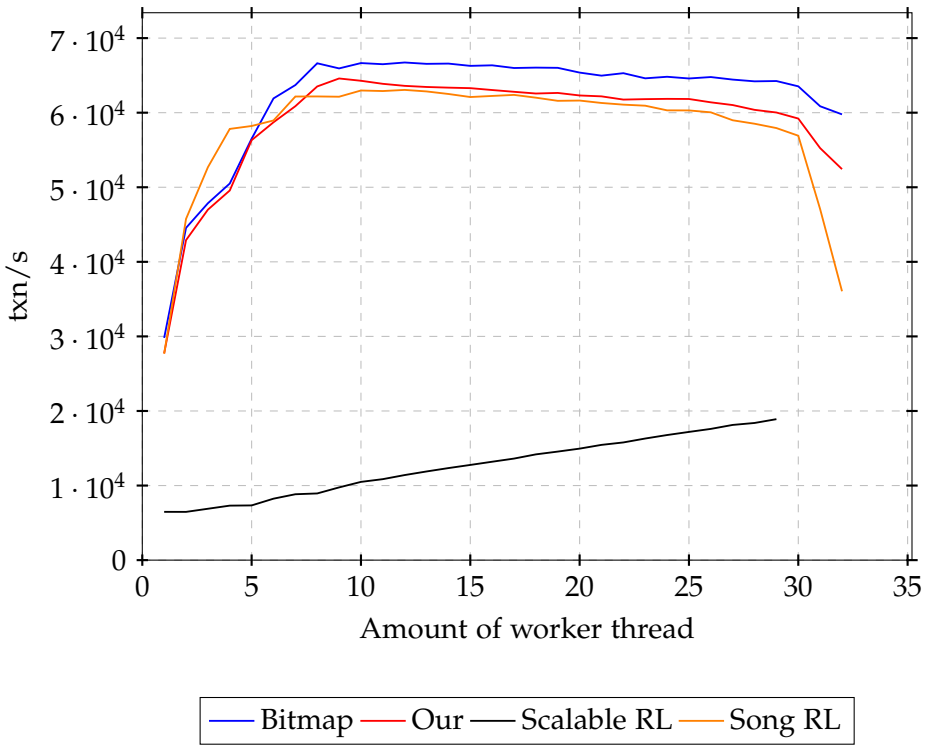


Figure 5.4: Leanstore YCSB: Transactional Throughput [Txn/s].

marked performance degradation. On the other hand, the `Scalable RL` showed steady, linear throughput growth. However, its overall performance remained considerably lower than the other competitors, highlighting its limitations.

An additional key observation is the difficulty faced by the `Bitmap`, `Our`, and `Song RL` competitors in scaling effectively beyond 16 worker threads. This limitation is due mainly to the BLOB size in our workload. With a BLOB size of 1MB and 16 worker threads, the combined size of the client-side buffer and the internal DBMS memory block for the BLOB exceeds the L3 cache capacity (32MB in our machine), resulting in significant contention at the L3 cache. Furthermore, LeanStore's use of `memcpy` for read operations exacerbates this issue. `memcpy` consumes 2MB of memory read and write for every 1MB BLOB. An average throughput of approximately 60,000 operations per second equals over 110 GB of memory consumed by `memcpy`, which saturates the memory bandwidth and hinders the application's ability to scale.

Metric

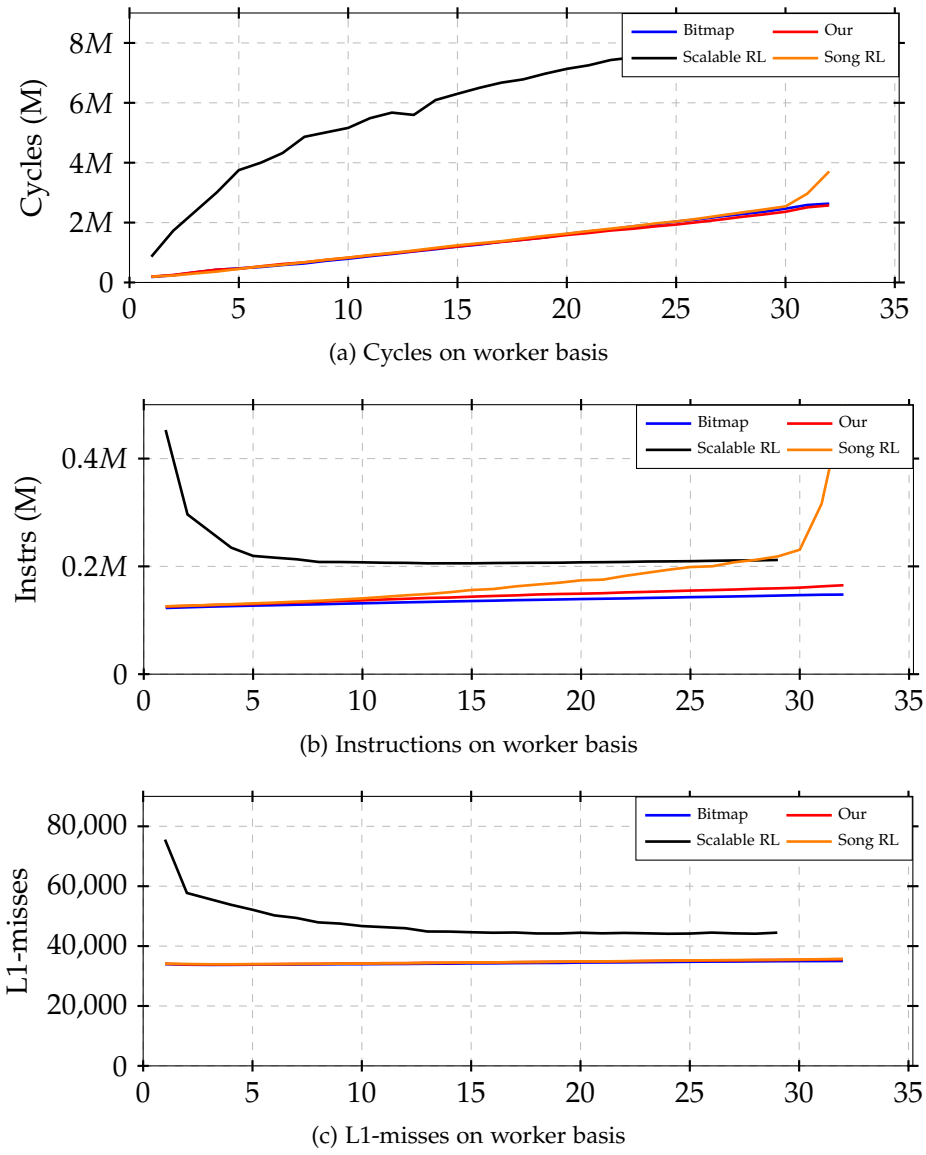
We also analyzed other performance of four methods in terms of CPU cycles, instructions, and L1 cache misses, as the number of workers increases. The result is depicted in Figure 5.4.

In the first graph, which measures cycles, the `Scalable RL` has the highest amount of cycles as the worker count rises. This is predictable as we have observed that `Scalable RL` performs poorly in the benchmark above. In contrast, `Bitmap`, `Song RL`, and `Our` methods maintain much lower and more stable cycle counts, with the `Our` method closely matching the `Bitmap` method in performance.

The second graph shows the number of instructions executed. The data indicates that, while the initial instruction count for `Scalable RL` is relatively high, it subsequently declines and stabilizes. The number of instructions of `Song RL` method exhibited a significant increase after reaching 25 threads, indicating the presence of a point of contention. Otherwise, `Our` method exhibits a slightly higher instruction count compared to `Bitmap`, but the difference remains small.

The third graph tracks L1 cache misses. It shows that `Scalable RL` suffers from the highest number of L1 misses, but again it subsequently declines and stabilizes. In contrast, `Bitmap`, `Song RL`, and `Our` methods keep their L1-miss counts low and stable across all workers.

Overall, the `Our` method achieves efficient resource usage, closely matching `Bitmap` in cycles and L1 misses, while showing a small increase in instructions. This suggests that the `Our` method is a scalable and efficient solution that outperforms `Scalable RL` and performs nearly as well as `Bitmap` across all metrics.



(a) Cycles on worker basis

(b) Instructions on worker basis

(c) L1-misses on worker basis

Figure 5.5: Cycles, instructions, and L1-misses on worker basis

6 Conclusion

This thesis introduced a new lock-free concurrent range lock aimed at overcoming the limitations of existing range-locking methods. Traditional techniques, like coarse-grained locks and tree or list-based locks, often struggle in high-performance computing because of contention and poor scalability under heavy loads. We proposed a new solution based on a probabilistic skip list to tackle these issues. By using atomic operations like `compareAndSwap` to manage node references, we avoided the need for traditional locks. This design reduces bottlenecks and improves performance, making it a better fit for large-scale distributed systems.

Our evaluation, which included detailed tests and integration into the Leanstore storage engine, showed that our approach outperforms existing range-locking methods. We saw significant improvements in throughput and reduced contention across different workloads and thread counts, confirming that our mechanism is effective in real-world high-concurrency scenarios.

Future work

Although we made significant progress, there are still areas that could be improved. One challenge we identified is that while our implementation works well, it doesn't scale as effectively when handling workloads beyond L3 cache sizes. This suggests a need to optimize our design for better cache efficiency, which would improve scalability in high-performance environments.

Another important area for future work is improving how we manage memory deallocation. The current design lacks an efficient memory cleanup strategy, which could lead to memory leaks or inefficient use of memory over time. Solving this issue will be key to enhancing the overall reliability and performance of our range-locking mechanism.

We also plan to take a closer look at the memory overhead created by our implementation. Using extra layers of linked lists and abstractions like `AtomicMarkableReference` adds some memory overhead. By analyzing this aspect and comparing it to other meth-

ods, we can better understand the trade-offs between memory usage and performance, which will help us refine our design.

Lastly, we want to explore how our concurrent range lock can be applied to other data structures or used in different contexts. For example, looking at the techniques used in the lock tree implementation in PerconaFT [22] could reveal new ways to improve the flexibility and usability of our approach. Investigating these areas could lead to further improvements and expand the potential applications of our range lock in high-concurrency environments.

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