

Partition-based Simple Heaps

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Abstract

We introduce a new family of priority-queue data structures: *partition-based simple heaps*. The structures consist of $\mathcal{O}(\lg n)$ doubly-linked lists; order is enforced among data in different lists, but the individual lists are unordered. Our structures have amortized $\mathcal{O}(\lg n)$ time extract-min and amortized $\mathcal{O}(\lg \lg n)$ time insert and decrease-key.

The structures require nothing beyond binary search over $\mathcal{O}(\lg n)$ elements, as well as binary partitions and concatenations of linked lists in natural ways as the linked lists get too big or small. We present three different ways that these lists can be maintained in order to obtain the stated amortized running times.

1. Introduction

The *priority queue* is a fundamental comparison-based abstract data type used in numerous efficient algorithms, such as discrete event simulation, Dijkstra’s shortest path algorithm and Prim’s minimum spanning tree algorithm. A priority queue (a.k.a. *heap*) stores a set of elements and has to implement the following operations, where we assume each element has an associated key from a comparison-based universe:

- **INSERT(e)**: Add an element e to the the heap and return a pointer ptr to it.
- **DELETE-MIN()**: Remove and return the element in heap with smallest key.
- **DECREASE-KEY(ptr, key)**: Change the key of the element pointed to by ptr in the heap to a smaller key key .

A standard undergraduate computer-science curriculum discusses binary heaps [Wil64], which realize all operations in $\mathcal{O}(\lg n)$ time¹ when the data structure currently contains n elements (as would a balanced binary search tree). However, much faster updates for priority queues are possible: The most efficient heaps such as (strict) Fibonacci heaps [FT87, BLT25] achieve *constant* (amortized) time for INSERT and DECREASE-KEY and $\mathcal{O}(\lg n)$ time for DELETE-MIN.

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¹Here and throughout we let \lg denote the binary logarithm.

While optimal in theory, these *efficient priority queues* have some downsides. First, their structure and analysis are less intuitive to teach and implement. Second, the constant factors in their running time and space usage are substantially higher than for binary heaps; not least because they inherently rely on pointer-based representations of trees, with the ensuing poor locality of reference.

The desire to simplify efficient priority queues and to make them more efficient in practice has resulted in a number of alternative data structures, which we review briefly in Section 1.1. They all retain somewhat complicated structures and invariants, are nontrivial to analyze, or require $\Omega(\lg n)$ time for updates.

In this paper, we present simple priority-queue implementations solely based on linear data structures, showing that simple quicksort-style partitioning of elements into unsorted sets suffices for amortized $\mathcal{O}(\lg \lg n)$ time INSERT and DECREASE-KEY operations and amortized $\mathcal{O}(\lg n)$ time DELETE-MIN operations – provided it is governed by a carefully chosen “*pivot-forgetting rule*” to keep the number of sets in $\mathcal{O}(\lg n)$. We present several such rules, demonstrating a rich design space. This brings us much closer to the efficient priority queues above, but with entirely elementary techniques and simple designs. As a case in point, for the arguably simplest rule (“lazy partition heaps”), we give a fully self-contained description including its amortized analysis suitable for teaching in undergraduate algorithms classes in Sections 2 and 3.

1.1. Previous Work

See [Bro13] for a full survey of priority queues. We begin the history of priority queues with fast DECREASE-KEY which begins with the Fibonacci heap [FT87]. They support constant-amortized time DECREASE-KEY and INSERT, and $\mathcal{O}(\lg n)$ amortized-time DELETE-MIN. Since then a variety of other heaps with fast DECREASE-KEY have been introduced:

- The pairing heap [FSST86], which is a self-adjusting structure modeled after splay trees, and which we do not know the asymptotic amortized running time of DECREASE-KEY. Only $\mathcal{O}(2^{\sqrt{\lg \lg n}})$ [Pet05] and $\Omega(\lg \lg n)$ [Fre99] are known. A multipass variant has been shown to have an $\mathcal{O}(\lg \lg n \cdot \lg \lg \lg n)$ amortized-time DECREASE-KEY [ST23a].
- The strict Fibonacci heap, which has worst-case running times [BLT25].
- Chan’s quake heaps, invented to be a simpler alternative to Fibonacci heap with $\mathcal{O}(1)$ amortized time DECREASE-KEY [Cha13].
- Rank-pairing heaps [HST09] which are another alternative with $\mathcal{O}(1)$ amortized time DECREASE-KEY, but where the structure is closer to the pairing heap.
- Violation heaps [Elm10], hollow heaps [HKTZ17], and Fibonacci heaps revisited [KTZ14] are other variants with with $\mathcal{O}(1)$ amortized-time DECREASE-KEY.
- Slim and smooth heaps [ST23b] have $\mathcal{O}(\lg \lg n)$ DECREASE-KEY, but are in the pointer-based model with constant indegree.

1.1.1. Lower bounds

There are two lower bounds for DECREASE-KEY in heaps, one due to Fredman [Fre99] and the other due to Iacono and Özkan [IÖ14]. Both of these bounds only apply to heap-based priority queues that follow two different particular models, each of which includes pairing heaps, and neither of which applies to our new structures. The Fredman lower bound focuses

on the tradeoff between augmented data and the decrease-key time. The Iacono and Özkan bound focuses on structures that are pointer-based. Both of these results point to the fact that $\mathcal{O}(\lg \lg n)$ is arguably a natural time for heaps, but both only apply to (different) classes of generalized heaps which are quite specific for their adversary arguments to work.

All of the above points to $\mathcal{O}(\lg \lg n)$ as the right running time for DECREASE-KEY in a pointer model structure with constant indegree. Slim and smooth heaps achieve this, and pairing heaps are conjectured to. All of the other heaps require either RAM-model tricks or non-constant indegree. For example, Fibonacci heaps need random access into an array of size $\mathcal{O}(\lg n)$ in order to combine heaps with the same rank (which is an integer with logarithmic range). One could of course replace this array with a balanced binary search tree, but this would introduce an $\mathcal{O}(\log \log n)$ overhead. In [FT87], it is shown how this array can be removed, albeit by using nodes of linear indegree.

1.1.2. Cache-oblivious inspiration

We point to the works of Chowdhury and Ramachandran [CR18] and Brodal, Fagerberg, Meyer, and Zeh [BFMZ04] as sources of inspiration as they are devoid of treeology. Both of these, which were designed for the cache-oblivious model, use two sequences of arrays of increasing size; superficially it looks like two copies of our data structure where items are inserted into one set, percolate down, at some point move to the other set, and then percolate up. Being designed for the cache-oblivious model, the lists are stored in arrays.

1.1.3. Lazy search trees

Also related are lazy search trees [SW20, SZ22, RW25], a data structure that smoothly interpolates between priority queues and binary search trees; depending on what queries are used, INSERT can be as fast as in the best priority queues, but arbitrary sorted dictionary queries are supported, as well. Our structures can be seen as a one-sided instance of the lower two tiers of the original lazy search trees with slightly modified operations.

1.1.4. Quickheaps

Other related work takes a more application-focused view of making priority queue implementations fast in practice. While asymptotic performance matters, constant-factor overhead and locality of reference can, for all reasonable input sizes, dwarf the distinction between logarithmic and sublogarithmic complexities. Quickheaps [PN06, NP10] and their “stronger” refinements [NPPS11] keep data in a single array and conceptually apply successive quickselect invocations for queries, where past partitioning steps are remembered and skipped.

This works well if updates come in random order; for robustness against adversarial inputs, stronger quickheaps take inspiration from schemes for balancing binary search trees to ensure good amortized performance in quickheaps: either using randomization or using a form of weight-balanced binary search tree. Being designed for the much harder problem of keeping an entire binary search tree in shape, the resulting concatenation rules for quickheaps are complicated. We point out that not only the concatenation rule itself, but also the amortized analysis in [NPPS11] is substantially different from those in this paper. They also cite favorable performance in the external-memory model as a benefit of quickheaps (in any variant). Our heaps can also have these same favorable properties. Note that the authors of quickheaps excluded pointer-based operations, in particular DECREASE-KEY, from their consideration of quickheaps in external memory.

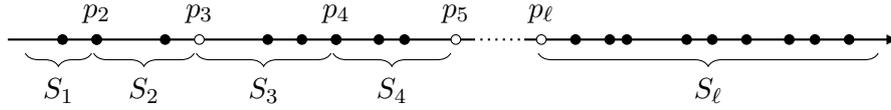


Figure 1: Illustration of our partition-based heap framework. The pivots $p_2 \dots, p_\ell$ partition the n elements in the heap into sets ℓ sets S_1, S_2, \dots, S_ℓ . Black dots are the n elements in the heap. Pivots can be elements in the heap (black) or elements previously removed from the heap (white).

2. Partition-based Heap Framework

Here we describe the generic framework of partition-based heaps common to all three of our data structures in Sections 3–5. We assume that the elements have keys from a totally ordered universe and that elements are ordered with respect to their keys. If ever duplicate keys are inserted into a partition-based heap, standard techniques are used so that all ties are broken consistently.

Partition-based heaps maintain the elements currently stored in the heap as a partition of sets S_1, S_2, \dots, S_ℓ , for some ℓ that is $\mathcal{O}(\lg n)$. The number of elements currently stored in the heap is $n = |S_1| + \dots + |S_\ell|$. While the sets themselves are unordered, if $i < j$, all elements in S_i are less than all elements of S_j .

Each set S_i is stored as a *doubly-linked list* as constant-time concatenation of sets is vital to all of our structures. For each set, the size $|S_i|$ is explicitly maintained. As we store all elements in linked lists, we support *stable pointers* to any element inserted.

In addition to the sets and their sizes, a partition-based heap stores *pivots* $p_2 \leq \dots \leq p_\ell$ which are elements such that $S_i \subseteq [p_i, p_{i+1})$ for $i = 1, \dots, \ell$; for notational convenience, assume p_1 and $p_{\ell+1}$ are pivots with keys $-\infty$ and $+\infty$, respectively. Pivots may be either elements currently or previously in the heap. See Figure 1 for an illustration.

To implement each priority queue operation all of our partition-based heaps have the following steps in common:

- **INSERT(e):** Search for e among the $\mathcal{O}(\lg n)$ pivots in $\mathcal{O}(\lg \lg n)$ time to find the set S_i with $p_i \leq e < p_{i+1}$. Insert e into S_i and increment the size $|S_i|$. Return a pointer ptr to the new node containing e in the linked list for S_i .
- **DELETE-MIN():** Remove the smallest element from the first nonempty set S_i , and decrement the size $|S_i|$.
- **DECREASE-KEY(ptr, key):** Let e denote the element in the linked-list node ptr points to. Remove this node from its linked list. Binary search on the pivots in $\mathcal{O}(\lg \lg n)$ to find the set S_i that the node was removed from and decrement the size $|S_i|$. Decrease the key of e to key , and re-insert the node as in INSERT, using a second search among the pivots to discover which set to insert it into, insert it at the end of the linked list for the set, and increment the set size.

These descriptions are not enough to maintain the crucial invariant that $\ell = \mathcal{O}(\log n)$. This is where the three heaps in Sections 3–5 differ; each has its own invariants towards this goal, and each augments the basic implementations above with restructuring operations which maintain said invariants. All structures use the technique of combining sets in constant time by concatenating linked-lists and splitting sets in linear time using a linear time selection algorithm

to maintain their invariants. Further, all structures use amortized analysis to argue for the running times.

At a very high level, the philosophy of each structure is as follows: In *lazy partition heaps* (LP heaps) in Section 3, the sets are allowed to become arbitrarily large, and there is no lower bound on the number of sets. A bound on the size of a set compared to the sum of all previous sets is maintained. In *Fibonacci heaps: the next generation* (FH:TNG) in Section 4 and the heaps in Section 5, sets are allowed to be empty and there are strict upper bounds on their size. In FH:TNG, there are lower bounds on non-empty set sizes, where both the upper and lower bounds are given by the Fibonacci numbers, which naturally is amiable to concatenating and splitting while maintaining the size invariants. The heaps in Section 5 do not maintain a lower bound on the set sizes which results in simpler restructuring algorithms.

3. Lazy Partition Heaps

To keep the number of sets ℓ in $\mathcal{O}(\lg n)$ at all times, *lazy partition heaps* (LP heaps) use the following pivot-forgetting rule, which is adapted from lazy search trees [SW20]:

A pivot p shall be abandoned if the two sets that it separates together contain fewer elements than there are smaller elements in the heap (Figure 2).

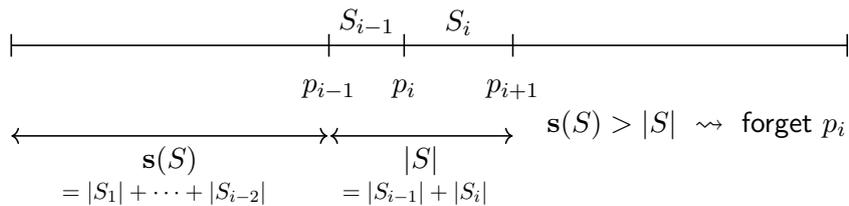


Figure 2: The pivot-forgetting rule for sets in LP heaps. Here, the lists for sets S_{i-1} and S_i would be concatenated (thus forgetting p_i) since there are more elements smaller than p_{i-1} than between p_{i-1} and p_{i+1} .

In LP heaps, every DELETE-MIN operation partitions S_1 around its median. Since no stringent bound on $|S_1|$ is guaranteed, we need to amortize the occasional high costs for DELETE-MIN. For that, sets S_j , $j = 1, \dots, \ell$, accumulate potential when they have grown “too large”, namely $\varphi_j := \max\{0, |S_j| - \mathbf{s}(S_j)\}$, where $\mathbf{s}(S_j) := |S_1| + \dots + |S_{j-1}|$ denotes the number of elements smaller than all elements in S_j (their “safety buffer” from the minimum).

This in particular includes the set S_1 , whose potential $\varphi_1 = |S_1|$ can thus pay for the scanning and partitioning of S_1 upon the next DELETE-MIN.

Apart from the sets, we store a (static) sorted array T of the current sets S_1, S_2, \dots, S_ℓ (in this order). (Whenever the collection of sets changes because of a DELETE-MIN, we can afford to rebuild T from scratch.) For simplicity of the pseudocode, we let each set S_i store the corresponding pivot p_i as $S_i.p$. We point out that $\mathbf{s}(S_j)$ is only needed during the pivot-forgetting phase and is computed on demand there.

3.1. Operations

Pseudo code for the operations on a LP heap is shown in Figure 3. Throughout the lifetime of an LP heap, we maintain the following invariant.

Invariant (S): The number of set is at most $\ell \leq 2 \lg n + 1$.

If the invariant is at risk of violation—which is only possible upon DELETE-MIN—we apply the following concatenation rule (cf. Figure 2). Recall that $\mathbf{s}(S_j) := |S_1| + \dots + |S_{j-1}|$ is the number of *smaller elements* (smaller than p_j , that is):

Concatenation Rule (C): If $|S_j| + |S_{j+1}| < \mathbf{s}(S_j)$, concatenate S_j and S_{j+1} .

Concatenation here simply means we concatenate the linked lists of elements for the two sets and “forgets” the pivot formerly separating the two sets. This operation takes constant time. FORGET-PIVOTS computes $\mathbf{s}(S_j)$ and applies rule (C) to \mathcal{S} successively. Clearly, FORGET-PIVOTS uses $\mathcal{O}(\ell)$ time if the (initial) number of sets is ℓ . Note that FORGET-PIVOTS also removes all empty sets present, thereby ensuring that $|S_j| \geq 1$ for all $1 \leq j \leq \ell$. After FORGET-PIVOTS has concatenated all pairs of sets that satisfy (C), we obtain a bound on the number of sets from the following lemma.

Lemma 3.1 ((C) implies (S)): *Suppose after applying rule (C), we have sets S_1, \dots, S_ℓ , with $|S_j| \geq 1$ for all $1 \leq j \leq \ell$. Then $\ell \leq 2 \lg n + 1$, for n the total size of the sets.* \triangleleft

Proof: We first show that prefixes of sets grow exponentially: for all $1 \leq j \leq \ell$, we have $\Sigma(j) := |S_1| + \dots + |S_j| \geq 2^{(j-1)/2}$ (*).

The proof is by induction on j . The claim holds for $1 \leq j \leq 2$, since $|S_i| \geq 1$ for $1 \leq i \leq \ell$. In the inductive step for $j \geq 3$, the inductive hypothesis implies $\Sigma(j-2) \geq 2^{(j-3)/2}$. By concatenation rule (C), we have $|S_{j-1}| + |S_j| \geq \mathbf{s}(S_{j-1})$, and by definition $\mathbf{s}(S_{j-1}) = \Sigma(j-2)$. Therefore $\Sigma(j) = \Sigma(j-2) + |S_{j-1}| + |S_j| \geq 2 \cdot \Sigma(j-2) \geq 2 \cdot 2^{(j-3)/2} = 2^{(j-1)/2}$, which proves (*).

Since (*) holds for any prefix of sets, it does so in particular for $j = \ell$. Then, $\Sigma(\ell) = n$ (all elements), and we get $\ell \leq 2 \lg n + 1$ as claimed. \square

It follows from Lemma 3.1 that FORGET-PIVOTS ensures (S). This completes the description of the data structure.

3.2. Potential

For analyzing the amortized cost of operations, we use the following potential:

$$\Phi := \sum_{j=1}^{\ell} \varphi_j \quad \text{with} \quad \varphi_j := \beta \cdot \max\{0, |S_j| - \mathbf{s}(S_j)\}.$$

(The sum is over all sets currently in the data structure.) Here $\beta \geq 2$ is a fixed scaling parameter (chosen depending on the constant factor in the linear cost of partitioning).

3.3. Amortized Cost of Insert

Upon an INSERT operation, the set S_j containing e is located using binary search in T in $\mathcal{O}(\lg \lg n)$ time. Adding e to the linked list of S_j takes constant time.

In Φ , $\mathbf{s}(S_{j'})$ increases by 1 for all $j' > j$; this never increases the value of Φ . The size of S_j grows by 1 which increases Φ by at most $\beta = \mathcal{O}(1)$. INSERT thus costs $\mathcal{O}(\lg \lg n)$ (both amortized and worst case).

LP-HEAP-INSERT(e)

```

1   $ptr := \text{NEW-NODE}(e)$  // obtain pointer to element
2   $S := T.\text{FIND}(e.\text{key})$  // binary search the pivots
3   $S.\text{APPEND}(ptr)$  // insert element
4  return  $ptr$ 

```

LP-HEAP-DECREASE-KEY(ptr, key)

```

1   $e := *ptr$  // obtain element from pointer
2   $S := T.\text{FIND}(e.\text{key})$ 
3   $S.\text{REMOVE}(ptr)$  // remove element from current set
4   $e.\text{key} := key$  // update key
5   $S := T.\text{FIND}(e.\text{key})$ 
6   $S.\text{APPEND}(ptr)$  // insert element in new set

```

LP-HEAP-DELETE-MIN()

```

1   $min := S_1.\text{DELETE-MIN}()$  // scans  $S_1$ 
2   $p := \text{MEDIAN}(S_1)$  // (larger) median (deterministic linear-time selection [BFP+73])
3   $\bar{S}_1, \bar{S}_2 := \text{PARTITION}(S_1, p)$ 
4   $\bar{S}_2.p := p$ 
5   $\mathcal{S} := [\bar{S}_1, \bar{S}_2, S_2, \dots, S_\ell]$  // doubly-linked list of sets
6   $\text{FORGET-PIVOTS}(\mathcal{S})$  // modifies  $\mathcal{S}$ 
7  Rebuild  $T$  from  $\mathcal{S}$ 
8  return  $min$ 

```

FORGET-PIVOTS(\mathcal{S})

```

// Assume  $\mathcal{S} = [S_1, \dots, S_\ell]$ 
1   $\mathcal{S}.\text{REMOVE-ALL-EMPTY-SETS}()$ 
2   $A := S_1$ 
3   $s(A) := 0$ 
4  while  $\mathcal{S}.\text{HAS-NEXT}(A)$ 
5       $B := \mathcal{S}.\text{NEXT}(A)$ 
6      if  $|A| + |B| < s(A)$ 
7          Concatenate  $B$  into  $A$ 
8           $\mathcal{S}.\text{REMOVE}(B)$ 
9      else
10          $s(B) := s(A) + |A|$ 
11          $A := B$ 
12     end if
13 end while

```

Figure 3: LP heaps implementation.

3.4. Amortized Cost of Decrease-Key

When $\text{DECREASE-KEY}(ptr, key)$ is performed, the element e at pointer ptr has its key decreased to key . Re-inserting e costs $\mathcal{O}(\lg \lg n)$ (both amortized and worst case). This changes the size of at most two sets, with one increasing and the other decreasing, which increases the potential by at most β . It may also change $\mathbf{s}(S_{j'})$ for some sets $S_{j'}$, but since we move e from S_i to an S_j with $1 \leq j \leq i$ when we make the key of the element smaller, $\mathbf{s}(S_j)$ can only *increase*, which does not increase the potential. In total, DECREASE-KEY uses $\mathcal{O}(\lg \lg n + \beta) = \mathcal{O}(\lg \lg n)$ time (both amortized and worst case).

DECREASE-KEY may leave a set S_i empty, where $2 \leq i \leq \ell$, however, these may be removed upon a later FORGET-PIVOTS , without an increase in time or potential, as the number of sets remains bounded, maintaining invariant (S).

3.5. Amortized Cost of Delete-Min

A DELETE-MIN scans S_1 to find the minimum element e and removes it; moreover, we split S_1 by partitioning it around its median, resulting in two new sets \bar{S}_1 and \bar{S}_2 . For the partition, we use deterministic selection. Partitioning S_1 has overall actual cost of $\mathcal{O}(|S_1|)$.

Finally, we establish invariant (S) by calling FORGET-PIVOTS . By the invariant (S) on the number of sets, and since partitioning set S_1 creates two new sets, FORGET-PIVOTS uses amortized $\mathcal{O}(\lg n)$ actual time.

It remains to analyze the change in potential. Note first that the concatenation rule (C) is chosen precisely so that when concatenating S_{j+1} into S_j , the contributions φ_{j+1} and φ_j , both before and after the concatenation, are all zero, so concatenating does not change the potential at all. More specifically, assume that $|S_j| + |S_{j+1}| < \mathbf{s}(S_j)$, and a concatenation of S_j and S_{j+1} is performed to obtain the new set \bar{S}_j , by the concatenation rule (C). Then the potential is decreasing by φ_j and φ_{j+1} and increasing by the new potential $\bar{\varphi}_j$. By definition these potentials are

$$\begin{aligned} \varphi_j &= \beta \cdot \max\{0, |S_j| - \mathbf{s}(S_j)\} &= 0 \\ \varphi_{j+1} &= \beta \cdot \max\{0, |S_{j+1}| - (\mathbf{s}(S_j) + |S_j|)\} &= 0 \\ \bar{\varphi}_j &= \beta \cdot \max\{0, (|S_j| + |S_{j+1}|) - \mathbf{s}(S_j)\} &= 0. \end{aligned}$$

The change in potential $\Delta\Phi$ is therefore 0 upon a concatenation.

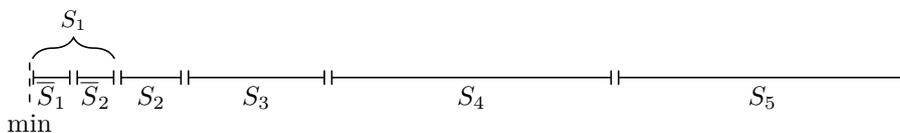


Figure 4: Sketch of the sets of an LP heap and the partition of S_1 into \bar{S}_1 and \bar{S}_2 in DELETE-MIN .

It remains to analyze the influence on the potential of partitioning the set S_1 into the new sets \bar{S}_1 and \bar{S}_2 . The situation after partitioning S_1 is illustrated in Figure 4. For the set S_1 , we lose $\varphi_1 = \beta|S_1|$ and gain $\bar{\varphi}_1 + \bar{\varphi}_2 = \beta|\bar{S}_1|$ from \bar{S}_1 and \bar{S}_2 , as we choose the larger median s. t. $\mathbf{s}(\bar{S}_2) = |\bar{S}_1| \geq |\bar{S}_2|$, i.e., $\bar{\varphi}_2 = 0$. All other sets lose one smaller element, so we gain $\leq \ell \cdot \beta$ in potential. By construction, $|\bar{S}_1| \leq \frac{1}{2}|S_1|$ (recall that we delete an element from S_1 before partitioning into \bar{S}_1 and \bar{S}_2), so $\Delta\Phi \leq -\frac{\beta}{2}|S_1| + \ell\beta$. For a sufficiently large $\beta \geq 2$, the total amortized cost is thus

$$\mathcal{O}(\Delta\Phi + |S_1| + \lg n) = \mathcal{O}(\lg n).$$

3.6. Extensions

Here we discuss a few further properties of LP heaps; they are not needed to understand (or teach) LP heaps.

3.6.1. Further Operations

We can support LP-HEAP-FIND-MIN in $\mathcal{O}(1)$ time, by maintaining the minimum upon LP-HEAP-INSERT and LP-HEAP-DECREASE-KEY in constant time, and by scanning \overline{S}_1 after a LP-HEAP-DELETE-MIN, which maintains the amortized time. Note that $|S_1|$ can be arbitrary large, i.e., we cannot afford scanning S_1 for each LP-HEAP-FIND-MIN operation.

We can also support an LP-HEAP-INCREASE-KEY operation or a general LP-HEAP-DELETE(*ptr*) operation, at total amortized cost $\mathcal{O}(\beta \lg n + \lg \lg n) = \mathcal{O}(\lg n)$. For both operations S_1 might become empty, and further for general LP-HEAP-DELETE(*ptr*), n decreases which may violate invariant (S), and therefore FORGET-PIVOTS is needed to re-establish (S). This also adds $\mathcal{O}(\lg n)$ to the cost.

An operation LP-HEAP-BUILD from a given collection of n elements can be supported in $\mathcal{O}(n)$ time. For that, we simply declare all elements in S_1 . The actual cost (if elements are given in a linked list) can even be constant, however we have to pay for the increase in potential of βn , so total amortized costs are $\mathcal{O}(n)$.

The LP-HEAP-MELD operation of taking the union of two heaps is not easily supported efficiently. One can trivially create a new LP heap from the concatenation of the two input heaps using LP-HEAP-BUILD, at linear amortized cost in the total number of elements.

3.6.2. Partitioning

We can replace the deterministic linear-time selection [BFP⁺73] by randomized selection [Hoa61, FR75], leading to the same expected amortized time.

More importantly, we can also simply run a *single round* of (random) partitioning: Upon the partition, the potential decreases by $\min\{|\overline{S}_1|, |\overline{S}_2|\}$. By choosing the pivot not as the median, but at random between the elements of S_1 , the decrease is therefore *expected* linear in $|S_1|$, and the amortized time of DELETE-MIN remains *expected* $\mathcal{O}(\lg n)$.

3.6.3. LP Quickheaps

The *quickheaps layout* (described in more detail in Section 1.1) offers an enticing alternative to the linked lists for each set: At the expense of more data movement during insertions and deletions, we store all sets contiguously in a single array, with the pivots separating them. The only extra space is a static array T with the $\mathcal{O}(\lg n)$ pivot positions.

Concatenations of adjacent sets can still be done in constant time (by simply demoting a pivot to an ordinary element), and partitioning can now be done in place, which improves practical overhead in running time. While we can also still binary search the pivots to *find* a set to insert an element into, insertions or deletions must now be implemented with up to one swap per set, leading to overall $\Theta(\lg n)$ running time for all operations.

3.6.4. Pointer-Machine Version

LP heaps already stay within the pointer-machine model of computation with bounded indegree nodes except for the index T . By replacing the sorted array and binary search by any balanced binary search tree, then LP heaps is entirely in the model.

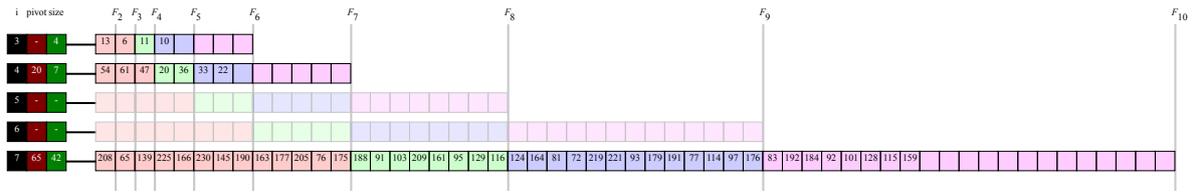


Figure 5: Illustration of a next generation Fibonacci heap. While each set is stored as a linked list, this illustration uses colors to show the relationship between each list's size, the Fibonacci numbers, and the size potential function.

4. Fibonacci Heaps: The Next Generation

The next generation Fibonacci heap (FH:TNG)² follows our framework for partition-based simple heaps with but a few distinctive features: Sets can either be present or absent. If a set is present, its size is required to be between the i th Fibonacci number, denoted as F_i ($F_1 := F_2 := 1$; $F_i := F_{i-1} + F_{i-2}$), and the $i + 3$ rd, F_{i+3} , with the one exception that the first set, S_3 , has no minimum size. Finally we require that there are at most eight empty sets in a row (*the consecutive empty invariant*) and at most two non-empty sets in a row (*the consecutive nonempty invariant*) to allow room for splits and concatenations.

The implementation of the operations follows our basic framework. The only difference is what to do when the invariants are violated. We restore them as follows using *invariant-restoring operations*: If a set S_i has F_{i+3} elements we call it *full* and there are two options: If the set below, S_{i+1} is empty, S_i is simply moved to S_{i+1} . We call this a *overflow-down operation*. If S_{i+1} is not empty, S_{i+2} must be empty by the consecutive nonempty invariant. We concatenate all F_{i+3} elements of S_i to the set S_{i+1} and then remove the F_{i+3} largest elements of the resultant set using linear-time selection and place them in S_{i+2} ; this leaves the size of S_{i+1} unchanged but the contents will certainly change. We call this an *overflow-thru operation*. If a set S_i has reached its minimum size F_i , it is *underfull*, and we execute either an *underflow-up operation* or an *underflow-thru operation* depending on whether S_{i-1} is empty. These are symmetric to the two overflow operations. If the consecutive-nonempty invariant is violated, there are three nonempty sets in a row with an empty list below. We concatenate the top two of them and place the resultant set in the empty slot below, we call this a *merge-down operation*.

If the consecutive-empty invariant is violated, there are nine empty sets in a row with a non-empty set below. We take the nonempty set below and split using linear-time selection into two set slots immediately above. We call this a *split-up*, and this takes linear time in the sets involved. Merge-down takes constant time as it is combining two linked lists, whereas the split-up takes linear time in the size of the sets to perform a selection and traverse the set to split it in two.

The asymmetry of the linear-time split-up with the constant time merge-down is crucial and is deeply connected to the fact that we allow keys to be decreased but not increased! The intuition is that repeated DECREASE-KEY operations can force inexpensive merge-downs to be performed as everyone is slowly moved to bigger sets. For example, this can happen by repeatedly DECREASE-KEY-ing the maximum item to make it the minimum. However, split-up operations are only triggered by the more expensive DELETE-MIN operations and thus their more expensive cost can be paid for.

²The name is meant to reflect that fact that in the original Fibonacci heaps and the the structure here, Fibonacci numbers play crucial, and entirely different, roles.

Here we present the potential function used in the analysis. The potential of the data structure is the sum of three components, each of which is defined for each set in the structure:

Nonempty potential. Define $\varphi_i^{\text{nonempty}}$ to be 1 if S_i is nonempty and 0 if S_i is empty. This is used to pay for the merge-down operation where the number of nonempty sets decreases by one at unit cost.

Size potential. Gives potential as S_i nears being full (F_{i+3} elements) or underfull (F_i elements) and 0 potential when its size is in the middle of its allowable range (from F_{i+1} to F_{i+2} elements). This is the classic potential function used in array doubling and is used to pay for the underflow and overflow operations. Figure 5 is color-coded to highlight these three ranges.

$$\varphi_i^{\text{size}} := \begin{cases} 0 & \text{if } S_i \text{ is empty} \\ F_{i+1} - |S_i| & \text{if } F_i \leq |S_i| < F_{i+1} \text{ (green in the figure)} \\ 0 & \text{if } F_{i+1} \leq |S_i| \leq F_{i+2} \text{ (blue in the figure)} \\ |S_i| - F_{i+2} & \text{if } F_{i+2} < |S_i| \leq F_{i+3} \text{ (purple in the figure)} \end{cases}$$

Up potential. This gives potential to nonempty sets that have few items above. This is used to pay for the split-up operation. The intuition is that this is a mechanism that allows the DELETE-MIN operation to pay unit cost to each of $\Theta(\lg n)$ sets, which the sets save up in the case most items above have been removed and a split-up needs to happen. Let \uparrow_i be defined to be $\sum_{j=1}^{i-1} |S_j|$. Define the up potential of set S_i , φ_i^\uparrow , as

$$\varphi_i^\uparrow := \begin{cases} 0 & \text{if } i \leq 2 \text{ or } S_i \text{ is empty} \\ \max\{0, F_{i-3} - \uparrow_i\} & \text{otherwise} \end{cases}$$

The potential of the data structure, Φ , is the sum of the up, size, and set potentials of every set; we refer to this sum as φ_i : $\Phi := \sum_{i=0}^k \overbrace{(\varphi_i^\uparrow + \varphi_i^{\text{size}} + \varphi_i^{\text{nonempty}})}^{\varphi_i}$.

Facts. We now present a series of facts which relating to these potentials.

Fact 4.1: If S_i is nonempty and at least of one of S_{i-1} , S_{i-2} S_{i-3} is nonempty then $\varphi_i^\uparrow = 0$. \triangleleft

Proof: If any of S_{i-1} , S_{i-2} and S_{i-3} are nonempty they will have least F_{i-3} items. Thus $\uparrow_i \geq F_{i-3}$ and thus $\max\{0, F_{i-3} - \uparrow_i\} = 0$. \square

Fact 4.2: $\uparrow_i \leq F_{i+4} - 1$ \triangleleft

Proof: $\uparrow_i = \sum_{j=1}^{i-1} |S_j| \leq \sum_{j=1}^{i-1} F_{i+3} = \sum_{j=4}^{i+2} F_i \leq F_{i+4} - 1$ via the classic identity that $\sum_{i=1}^n F_i = F_{i+2} - 1$ \square

Fact 4.3: If an operation redistributes items among S_i, S_{i+1}, \dots, S_j without insertion or deletion, φ_ℓ^\uparrow does not change outside of $\ell \in [i, j]$. \triangleleft

Proof: Recall that φ_ℓ^\uparrow is a function of whether S_ℓ is empty as well as \uparrow_ℓ , which is defined to be $\sum_{m=1}^{\ell-1} |S_m|$ and does not change if the sets being redistributed are either entirely inside or entirely outside of this sum. \square

Fact 4.4: For $i = O(1)$, $\varphi_i = O(1)$. ◁

The next fact, follows from the intuition that set S_i can be split into S_{i-1} and S_{i-2} or an inverse concatenation performed without gaining size potential. This follows from the properties of Fibonacci numbers, where a set of a certain fullness can be split into two with the same fullness and thus the same potential.

Fact 4.5: The elements of a nonempty set S_i can be distributed to previously empty sets S_{i-1} and S_{i-2} so that the set potential remains unchanged. Inversely, the elements of nonempty sets S_{i-1} and S_{i-2} can be concatenated into the previously empty set S_i without increasing potential. ◁

Proof: In general, if S_i has between F_{i+j} and F_{i+j+1} elements, for $j \in 0, 1, 2$ we can proportionally distribute them so that S_{i-1} has between F_{i+j-1} and F_{i+j} elements and S_{i-2} has between F_{i+j-2} and F_{i+j-1} elements; this follows naturally from the definition of Fibonacci numbers. In any such distribution, for the case where $j = 0$ the size potential was and remains, $F_{i-1} - |S_i|$, for $j = 1, 0$, and for $j = 2$, $|S_i| - F_{i+2}$.

For the inverse operation of merging, set j so that S_i will have between F_{i+j} and F_{i+j+1} elements. If $j = 1$ then this holds trivially as the potential will become zero. For the cases where $j = 0$ and $j = 2$ a little math confirms the intuition:

$$\underbrace{\varphi_i^{\text{size}}}_{\text{after}} = |S_i| - F_{i+1} = \overbrace{|S_{i-1}| - F_i + |S_{i-2}| - F_{i-1}}^{\substack{|S_{i-1}| - F_i \leq \varphi_{i-1}^{\text{size}} \\ |S_{i-2}| - F_{i-1} \leq \varphi_{i-2}^{\text{size}}}} \leq \underbrace{\varphi_{i-1}^{\text{size}} + \varphi_{i-2}^{\text{size}}}_{\text{before}}$$

$$\underbrace{\varphi_i^{\text{size}}}_{\text{after}} = F_{i+3} - |S_i| = \overbrace{F_{i+2} - |S_{i-1}| + F_{i+1} - |S_{i-2}|}^{\substack{F_{i+2} - |S_{i-1}| \leq \varphi_{i-1}^{\text{size}} \\ F_{i+1} - |S_{i-2}| \leq \varphi_{i-2}^{\text{size}}}} \leq \underbrace{\varphi_{i-1}^{\text{size}} + \varphi_{i-2}^{\text{size}}}_{\text{before}} \quad \square$$

4.1. Split-up

See Figure 6. Split-up is called when the invariant that there are at most 8 consecutive empty sets in a row is violated, that is, there are 9 consecutive empty sets. Specifically, it is called on set i when S_i and S_{i-10} are nonempty and $S_{i-1} \dots S_{i-9}$ are empty. This requires $i \geq 11$.

The moving of items from S_i to S_{i-1} and S_{i-2} should be done in a way proportional their sizes, that is in respect of the golden ratio. Very specifically, if there are between F_l and F_{l-1} elements in S_i , these should be moved so there are between F_{l-1} and F_{l-2} elements in S_{i-1} and F_{l-2} and F_{l-3} elements in S_{i-3} ; this is of course possible due to the definition of the Fibonacci numbers.

The new pivots are set to the minimum in each set via linear scan.

Lemma 4.6: *The amortized cost of the split-up operation is at most 0 when called on S_i , $i \geq 11$.* ◁

Proof: The nominal cost of the split-up operation is defined to be F_{i-6} as it takes linear time in the sets involved. To compute the change in potential we look at each of the three components of the potential function.

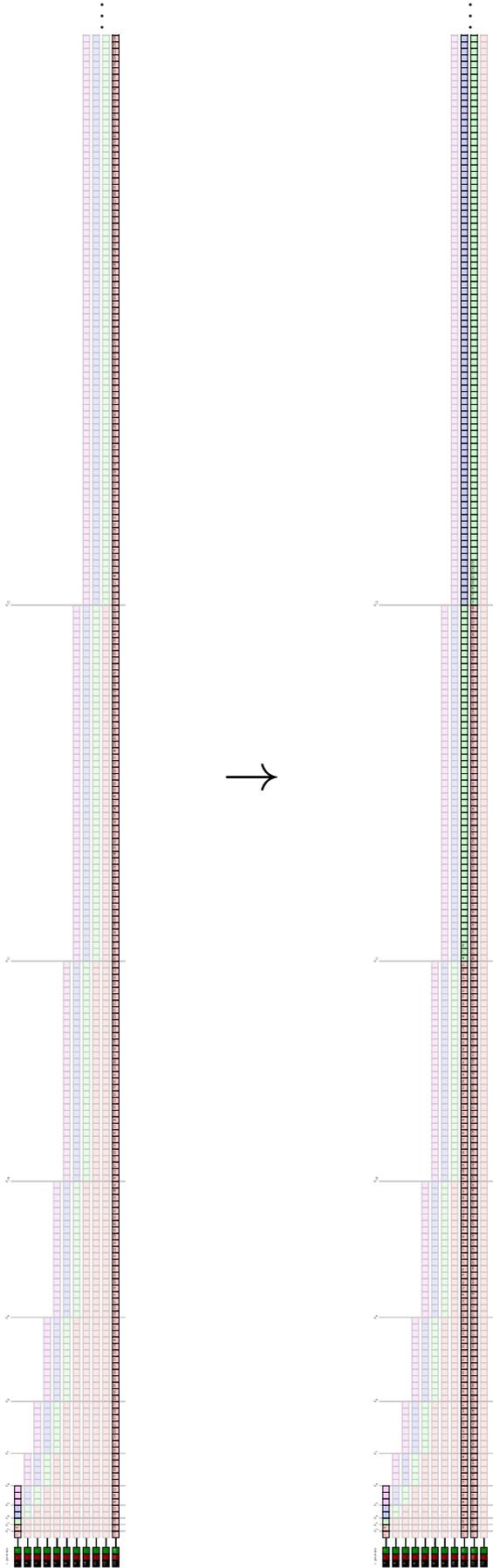


Figure 6: Split-up operation. There are 9 consecutive empty sets in a row, $S_2 \dots S_{10}$. The set S_{11} is then proportionally split into sets S_9 and S_{10} . In all cases the nonempty sets remain in the same color range, in this case green, which represents one of the three cases of the size potential. Empty cells continue off the page.

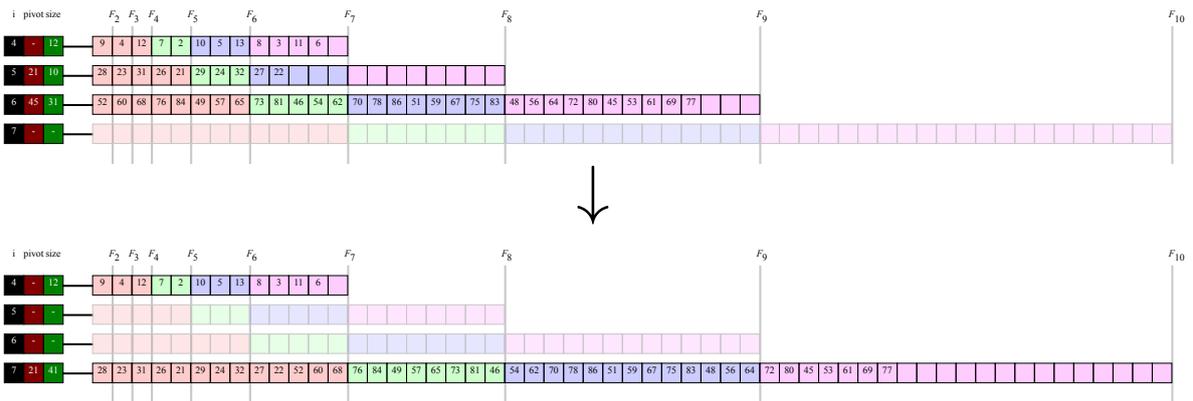


Figure 7: A merge-down operation performed when there are three non-empty sets in a row, here S_4, S_5, S_6 , and the set below S_7 is empty. sets S_5 and S_6 are concatenated into the empty slot of set S_7 .

Change in φ_i^\uparrow : This is $\sum_{i=0}^k \varphi_i^\uparrow - \sum_{i=0}^k \bar{\varphi}_i^\uparrow$. However, this can be simplified using Fact 4.3 and recalling that empty sets have zero up potential to $\bar{\varphi}_{i-1}^\uparrow + \bar{\varphi}_{i-2}^\uparrow - \varphi_i^\uparrow$. We can now bound each of these quantities separately. First, $\bar{\varphi}_{i-1}^\uparrow = 0$ by Fact 4.1. Next, for $\bar{\varphi}_{i-2}^\uparrow$, this is $\max\{0, F_{i-5} - \bar{\uparrow}_{i-2}\} \leq F_{i-5}$.

Finally, for $-\varphi_i^\uparrow$: we use the fact that $S_{i-1} \dots S_{i-9}$ are empty to bound \uparrow_i :

$$\begin{aligned}
-\varphi_i^\uparrow &= -\max\{0, F_{i-3} - \uparrow_i\} && \text{By definition} \\
&= -\max\{0, F_{i-3} - \uparrow_{i-9}\} && \uparrow_i = \uparrow_{i-9} \\
&< -\max\{0, F_{i-3} - F_{i-5} - 1\} \uparrow_{i-9} && < F_{i-5} - 1 \text{ by Fact 4.2} \\
&= -F_{i-4} + 1 && i \geq 11
\end{aligned}$$

Thus, bringing it together gives:

$$\sum_{i=0}^k \varphi_i^\uparrow - \sum_{i=0}^k \bar{\varphi}_i^\uparrow = \bar{\varphi}_{i-1}^\uparrow + \bar{\varphi}_{i-2}^\uparrow - \varphi_i^\uparrow \leq 0 + F_{i-5} - F_{i-4} + 1 = -F_{i-6} + 1$$

Change in φ_i^{size} : 0.

Change in $\varphi_i^{\text{nonempty}}$: +1, as there is now one more nonempty set.

In summary, by definition, the amortized cost is the nominal cost plus the change in potential which is at most, when $i \geq 11$: $F_{i-6} - F_{i-6} - 1 + 0 + 1 = 0$ \square

4.2. Merge-down

See Figure 7. The merge-down operation is called when the invariant that there are most two non-empty consecutive sets is violated. That is, there are three non-empty consecutive sets S_i, S_{i-1} and S_{i-2} with S_{i+1} being empty (or non-existent). The sets S_i and S_{i-1} are concatenated to form the new set \bar{S}_{i+1} ; this is where we use fact that the sets are stored as doubly-linked sets and not as arrays so that this operation takes constant time. The pivot \bar{p}_{i-2} is set to p_{i-1} .

Lemma 4.7: *The amortized cost of the merge-down operation is at most 0.* \triangleleft

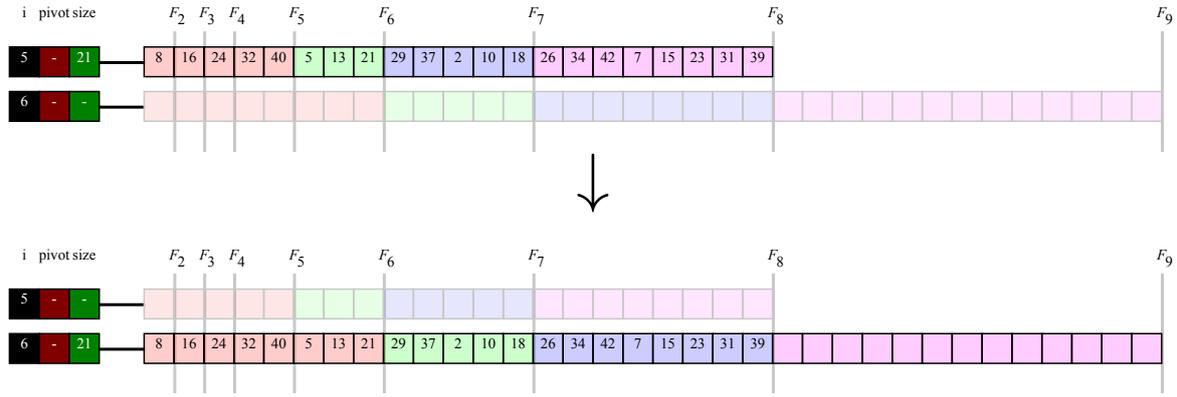


Figure 8: An overflow-down operation performed on S_5 which has $F_8 = 21$ elements when S_6 is empty.

Proof: The nominal cost of the operation is 1, To compute the change in potential we look at each of the three components of the potential function:

Change in φ_i^\uparrow : 0. By Fact 4.3 this is $\bar{\varphi}_{i+1}^\uparrow - \varphi_i^\uparrow - \varphi_{i-1}^\uparrow$. By Fact 4.1, since $S_{i-2} = \bar{S}_{i-2}$ is nonempty, all three of these are zero.

Change of φ_i^{size} : ≤ 0 by Fact 4.5.

Change of $\varphi_i^{\text{nonempty}}$: -1 as the number of non-empty sets decreases by one.

In summary, by definition, the amortized cost is the nominal cost plus the change in potential:
 $1 + 0 + 0 - 1 = 0$ □

4.3. Overflow-down

See Figure 8. The overflow-down operation is called when a set S_i is full, and thus has F_{i+3} items, and the set below, S_{i+1} , is empty. The set is moved from S_i to S_{i+1} in constant time.

Lemma 4.8: *The amortized cost of the overflow-down on set S_i is at most 0 when $i \geq 3$* ◁

Proof: The nominal cost of overflow up is 1, since it is a constant-time operation. To compute the change in potential we look at each of the three components of the potential function:

Change in φ_i^\uparrow : By Fact 4.3 this is $\bar{\varphi}_{i+1}^\uparrow - \varphi_i^\uparrow$ which is at most $\bar{\varphi}_{i+1}^\uparrow$. Using the definition of the up potential, $\bar{\varphi}_{i+1}^\uparrow = \max\{0, F_{i-2} - \uparrow_{i+1}\} \leq F_{i-2}$.

Change of φ^{size} : As $|\bar{S}_{i+1}| = F_{i+3}$, $\bar{\varphi}_{i+1}^{\text{size}} = 0$. Thus we need only consider $-\varphi_i^{\text{size}}$: $-\left(\overbrace{|S_i|}^{F_{i+3}} - F_{i+2}\right) = F_{i+2} - F_{i+3} = -F_{i+1}$

Change of $\varphi_i^{\text{nonempty}}$: 0 as the number of nonempty sets does not change.

In summary, by definition, the amortized cost is the nominal cost plus the change in potential
is $1 + \underbrace{F_{i-2} - F_{i+1}}_{\leq -F_i} + 0 \leq 0$. □

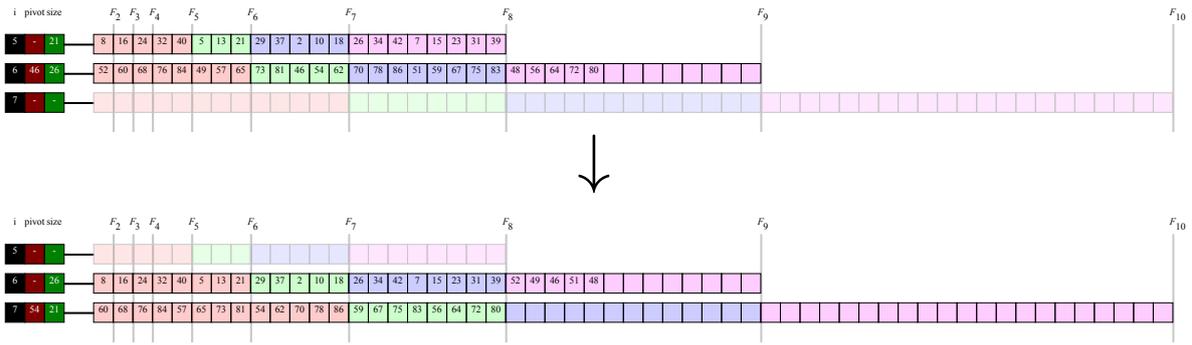


Figure 9: An overflow-thru operation performed on S_5 which has $F_8 = 21$ elements when S_6 is nonempty.

4.4. Overflow-thru

See Figure 9. The overflow-thru operation is called when a set S_i is full, and thus has F_{i+3} items, and the set below, S_{i+1} is nonempty. As there can not be more than two non-empty consecutive sets by the consecutive nonempty invariant, this implies S_{i+2} is empty.

Recall that in this operation, S_i and S_{i+1} are combined, and the largest items are moved from the combined set to \bar{S}_{i+2} so that the size of \bar{S}_{i+1} is unchanged and thus $|\bar{S}_{i+2}| = |S_i| = F_{i+3}$. This movement requires a linear time median/order statistic algorithm and thus has nominal cost F_{i-4} since S_{i+2} is the largest set involved. The pivot \bar{p}_{i+1} and \bar{p}_{i+2} are set to the minimum in each set.

Lemma 4.9: *The amortized cost of overflow-thru is ≤ 0 , when $i \geq 5$.* \triangleleft

Proof: The nominal cost of the operation is F_{i-4} as it takes linear time in the size of the sets involved. To compute the change in potential we look at each of the three components of the potential function:

Change in φ^\uparrow : By Fact 4.3 this is $\bar{\varphi}_{i+2}^\uparrow + \bar{\varphi}_{i+1}^\uparrow - \varphi_{i+1}^\uparrow - \varphi_i^\uparrow$. By Fact 4.1, φ_{i+1}^\uparrow and $\bar{\varphi}_{i+2}^\uparrow$ are both zero. Combined with the fact that φ_{i+1}^\uparrow is nonnegative, the change in φ^\uparrow is at most $\bar{\varphi}_{i+1}^\uparrow = \max\{0, F_{i-2} - \uparrow_{i+1}\} \leq F_{i-2}$.

Change in φ^{size} : This is $\bar{\varphi}_{i+2}^{\text{size}} + \bar{\varphi}_{i+1}^{\text{size}} - \varphi_{i+1}^{\text{size}} - \varphi_i^{\text{size}}$. As $|S_{i+2}| = F_{i+3}$, $\bar{\varphi}_{i+2}^\uparrow = 0$ by definition. Since $\varphi_{i+1}^{\text{size}} = \bar{\varphi}_{i+1}^{\text{size}}$, this means the change in φ^{size} is $-\varphi_i^{\text{size}}$. By definition $-\varphi_i^{\text{size}} = -(|S_i| - F_{i+2})$, which since $|S_i| = F_{i+3}$ is $-F_{i+1}$.

Change in $\varphi^{\text{nonempty}}$: 0, since the total number of nonempty sets does not change.

In summary, by definition, the amortized cost is the nominal cost plus the change in potential: $F_{i-4} + F_{i-2} - F_{i+1} + 0 \leq 0$ \square

4.5. Underflow-up

See Figure 10. The underflow-up operation is called when a set S_i is underfull, and thus has F_i items, and the set above, S_{i-1} , is empty. In this case the set and the pivot is moved from S_i to S_{i-1} in constant time.

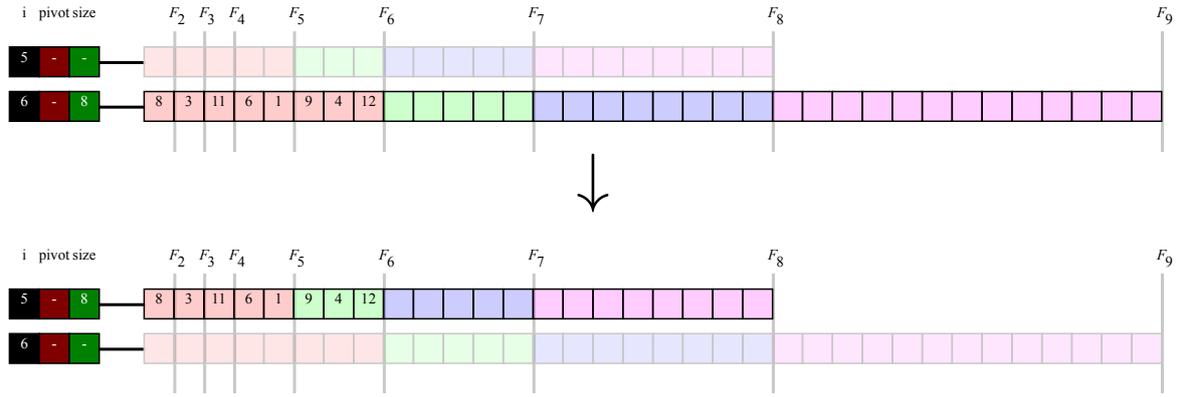


Figure 10: An underflow-up operation performed on S_6 which has $F_6 = 8$ elements when S_5 is empty.

Lemma 4.10: *The amortized cost of the underflow-up operation on S_i is at most 0 when $i \geq 6$.* ◁

Proof: The nominal cost of the operation is 1, as it is a constant-time operation. To compute the change in potential we look at each of the three components of the potential function:

Change in φ^\uparrow : By Fact 4.3 this is $\bar{\varphi}_{i-1}^\uparrow - \varphi_i^\uparrow$. We have $\bar{\varphi}_{i-1}^\uparrow - \varphi_i^\uparrow \leq \bar{\varphi}_{i-1}^\uparrow \leq F_{i-4}$ since $\bar{\varphi}_{i-1}^\uparrow = \max\{0, F_{i-4} - \bar{\tau}_{i-1}\}$ by definition.

Change in φ^{size} : This is $\bar{\varphi}_{i-1}^{\text{size}} - \varphi_i^{\text{size}}$. As $|\bar{S}_{i-1}| = F_i$, $\bar{\varphi}_{i-1}^{\text{size}} = 0$ by definition. We can now

$$\text{bound } -\varphi_i^{\text{size}}: -(F_{i+1} - \overbrace{|S_i|}^{F_i}) \leq -F_{i+1} + F_i = -F_{i-1}.$$

Change in $\varphi^{\text{nonempty}}$: 0. The number of nonempty sets does not change. ◻

In summary, by definition, the amortized cost is the nominal cost plus the change in potential: $1 + F_{i-4} - F_{i-1} + 0 < 0$ as $i \geq 6$.

4.6. Underflow-thru

See Figure 11. The underflow-thru operation is called when a set S_i is underfull, and thus has F_i items, and the set above, S_{i-1} is nonempty. As there can not be more than two non-empty consecutive sets by the consecutive nonempty invariant, this implies S_{i-2} is nonempty.

Recall that in this operation, S_i and S_{i-1} are combined, and the smallest items are moved from the combined set to \bar{S}_{i-2} so that the size of S_{i-1} is unchanged, thus $|\bar{S}_{i-1}| = |S_{i-1}|$ and $|\bar{S}_{i-2}| = |S_i| = F_i$. This requires a linear time median/order statistic algorithm and thus has nominal cost F_{i-6} since S_i is the largest set involved. The pivots \bar{p}_{i-1} and \bar{p}_{i-2} are set to the minimum in each set.

Lemma 4.11: *The amortized cost of the underflow-thru operation is at most 0 when $i \geq 3$.* ◁

Proof: The nominal cost of the operation is F_{i-6} . To compute the change in potential we look at each of the three components of the potential function:

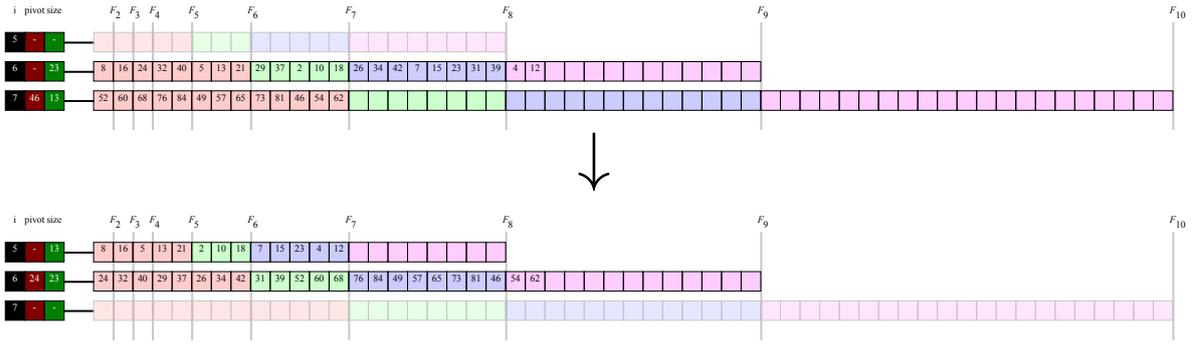


Figure 11: An underflow-thru operation performed on S_7 which has $F_7 = 13$ elements when S_5 is nonempty.

Change in φ^\uparrow : By Fact 4.3 this is $\overline{\varphi}_{i-1}^\uparrow + \overline{\varphi}_{i-2}^\uparrow - \varphi_i^\uparrow - \varphi_{i-1}^\uparrow$. By Fact 4.1, φ_i^\uparrow and $\overline{\varphi}_{i-1}^\uparrow$ are both zero. Also, $-\varphi_{i-1}^\uparrow$ is at most 0. Thus the change in φ^\uparrow is at most $\overline{\varphi}_{i-2}^\uparrow$ which is $\max\{0, F_{i-5} - \overline{\tau}_{i-2}\} \leq F_{i-5}$.

Change in φ^{size} : This is $\overline{\varphi}_{i-2}^{\text{size}} + \overline{\varphi}_{i-1}^{\text{size}} - \varphi_{i-1}^{\text{size}} - \varphi_i^{\text{size}}$. Note that while the contents of S_{i-1} will change, the size will not and thus $\varphi_{i-1}^{\text{size}} = \overline{\varphi}_{i-1}^{\text{size}}$. Also note that by definition, the size of \overline{S}_{i-2} is F_i and thus by definition $\overline{\varphi}_{i-2}^{\text{size}} = 0$. So, the change in φ^{size} is $-\varphi_i^{\text{size}}$:

$$-\varphi_i^{\text{size}} - (F_{i+1} - \overbrace{|S_i|}^{F_i}) \leq -(F_{i+1} - F_i) = -F_{i-1}$$

Change in $\varphi^{\text{nonempty}}$: 0 The number of nonempty sets does not change.

In summary, by definition, the amortized cost is the nominal cost plus the change in potential: $F_{i-6} + F_{i-5} - F_{i-1} + 0 < 0$. \square

4.7. Insert

In an INSERT operation, a binary search is performed on the pivots to determine which set the newly inserted item is added to. If the item is smaller than the smallest pivot, it is added to the first set, S_3 . Denote this set as j . The size of j is updated and the pivot of j is not updated as it is still valid. If this becomes overfull, a single overflow-down or overflow-thru is performed which eliminates the invariant violation on the set size but may cause too many empty sets in a row below j and too many nonempty sets in a row above j . The former is fixed via split-up, and the latter via merge-down, which can cascade downwards.

Lemma 4.12: *The amortized cost of the INSERT operation is $O(\lg \lg n)$.* \triangleleft

Proof: We need not analyze any invariant-restoring operations as these have been shown to have amortized cost at most zero (except for those called on constant sized sets which can happen at most a constant number of times with a constant amortized cost).

The nominal cost of this operation is $O(\lg \lg n)$ due to the binary search on the pivots of the $\Theta(\lg n)$ sets.

To compute the change in potential we look at each of the three components of the potential function.

Change in φ_i^\uparrow : As this is defined as $\max\{0, F_{i-3} - \uparrow_i\}$, and the \uparrow_i are only incremented if they change, this potential increases by at most 0.

Change in φ_i^{size} : ≤ 1 . The set j now has one more element and thus could gain one unit of size potential.

Change in $\varphi_i^{\text{nonempty}}$: 0, as the number of sets do not change.

In summary, by definition, the amortized cost excluding any invariant-restoring operations is the nominal cost plus the change in potential which is at most: $O(\lg \lg n) + 0 + 1 + 0 = O(\lg \lg n)$. Including any calls to invariant-restoring operations, which have a total of $O(1)$ amortized cost, gives the total amortized cost of $O(\lg \lg n)$. \square

4.8. Delete-Min

In the DELETE-MIN operation, we find and remove the smallest element from the smallest non-empty set, which we denote as S_j . Due to the consecutive empty invariant, $j \leq 11$ and thus this takes constant time as $|S_j| \leq F_{14} = 377$. We update the size of the set j . If the set becomes underfull, we perform a single underflow-up or underflow-thru, which can cause a cascading chain of merge-downs to be performed.

Lemma 4.13: *The amortized cost of the DELETE-MIN operation is $O(\lg n)$.* \triangleleft

Proof: We need not analyze any invariant-restoring suboperations as these have been shown to have amortized cost at most zero (except for those called on constant sized sets which can happen at most a constant number of times with a constant amortized cost).

The actual cost of this operation excluding the invariant-restoring sub-operations is $O(1)$.

To compute the change in potential we look at each of the three components of the potential function.

Change in φ_i^\uparrow : As this is defined as $\max\{0, F_{i-3} - \uparrow_i\}$ for the nonempty sets (and 0 for the empty sets), and each of the $O(\lg n)$ \uparrow_i will decrease by at most 1, the up potential increases by at most $O(\lg n)$.

Change in φ_i^{size} : ≤ 1 . The set j now has one less element and thus could gain one unit of size potential.

Change in $\varphi_i^{\text{nonempty}}$: At most 0.

In summary, by definition, the amortized cost is the actual cost plus the change in potential which is: $O(1) + O(\lg n) + 1 + 0 = O(\lg n)$. Including any calls to invariant-restoring operations, which have a total of $O(1)$ amortized cost, gives the total amortized cost of $O(\lg n)$. \square

4.9. Decrease-Key

This operation is carried out by first removing the item from its current set, S_j , decreasing it, and re-inserting it in the appropriate set S_k ; crucially $k \leq j$ as the key is only decreased. Two binary searches are needed on the previous value (to update the size of the set it was in) and the new set (to insert and also update the size). The pivots of both sets remain valid with no change needed. After the removal, if the set is underfull a single underflow-up or underflow-thru is performed, and then split-up and merge-downs are performed as needed until all invariants are restored; after the insertion if the set is overfull a single overflow-down or overflow-thru is performed, and then split-up and merge-down are performed as needed until all invariants are restored.

Lemma 4.14: *The amortized cost of DECREASE-KEY is $\mathcal{O}(\lg \lg n)$.* ◁

Proof: We need not analyze any invariant-restoring suboperations as these have been shown to have amortized cost at most zero (except for those called on constant sized sets which can happen at most a constant number of times with a constant amortized cost).

The nominal cost of this operation is $\mathcal{O}(\lg \lg n)$ due to the two binary search operations.

To compute the change in potential we look at each of the three components of the potential function.

Change in φ_i^\uparrow : This is defined as $\max\{0, F_{i-3} - \uparrow_i\}$ for the nonempty sets, and 0 for the empty sets. For $i < k$ and $i > j$ this will not change, but for i between k and j , that is for sets between the removal and re-insertion, \uparrow_i will increase by 1, and thus φ_i^\uparrow will decrease by 1. This is where we crucially use the fact that keys can only decrease, as if we allowed an increase the potential here could grow by $\Theta(\lg n)$, but instead it may only shrink.

Change in φ_i^{size} : ≤ 2 . The insertion and deletion could each cause a unit change in the size potential of the affected sets.

Change in $\varphi_i^{\text{nonempty}}$: 0, as the number of sets do not change.

In summary, by definition, the amortized cost is the actual cost plus the change in potential which is at most: $\mathcal{O}(\lg \lg n) + 0 + 2 + 0 = \mathcal{O}(\lg \lg n)$. Including any calls to invariant-restoring operations, which have a total of $\mathcal{O}(1)$ amortized cost, gives the total amortized cost of $\mathcal{O}(\lg \lg n)$. □

5. Heaps with Exponential Upper Bounded Sets

In this section we consider another variant of a partition-based simple heap consisting of sets S_1, \dots, S_ℓ and pivots $p_2 \leq \dots \leq p_\ell$. Again, the time obtained for DELETE-MIN is amortized $\mathcal{O}(\lg n)$ and for INSERT and DECREASE-KEY amortized $\mathcal{O}(\lg \lg n)$. In this variant we have exponential increasing upper bounds on the sizes of the sets, without any lower bounds, and allowing an arbitrary number of the sets to be empty. This is in contrast to the stronger structural invariants of the FH:TNG in Section 4.

Invariant: $|S_i| < 3 \cdot 2^i$ for $1 \leq i \leq \ell$.

The heap operations are implemented as follows:

- INSERT(e):
Binary search over the pivots to find i where $p_i \leq e < p_{i+1}$, and append e to S_i . If $|S_i| = 3 \cdot 2^i$, perform a recursive *push* of S_i to S_{i+1} (see below).
- DELETE-MIN():
If S_1 is empty, first recursively *pull* elements into S_1 from S_2 (see below). Delete and return the minimum from S_1 . If $\ell > 1 + \lg n$, let $S_{\ell-1}$ be the concatenation of $S_{\ell-1}$ and S_ℓ , and discard S_ℓ and p_ℓ before returning, i.e., ℓ decreases by one (this ensures $\ell \leq 1 + \lg n$).
- DECREASE-KEY(ptr , key):
Delete the element e pointed to by ptr from its current set S_j (to decrement the size $|S_j|$ we find j by a binary search over the pivots). Change the key of e to key , and perform a binary search over the pivots to find the set S_i where to add e , and append e to S_i . If $|S_i| = 3 \cdot 2^i$, perform a recursive *push* of S_i to S_{i+1} .

Push. For INSERT and DECREASE-KEY, the set S_i receiving one more element might get size $|S_i| = 3 \cdot 2^i$. We *push* all $3 \cdot 2^i$ elements of S_i to S_{i+1} , and set $S_i = \emptyset$. In general, a recursive push to S_j of a set of elements X always satisfies $2^{j-1} \leq |X| \leq 3 \cdot 2^{j-1}$. If $|S_j| < 2^j$ before the push (i.e., S_j is too small to be pushed to S_{j+1}), X is concatenated to S_j in constant time, such that $|S_j| < 5 \cdot 2^{j-1}$, and the push terminates. Otherwise, the set X becomes S_j and the old S_j is recursively pushed to S_{j+1} . The pivots need to be updated accordingly, and when $j = \ell + 1$ a new set $S_{\ell+1} = X$ is created, i.e., ℓ increases by one.

Pull. The recursive *pull* of elements into S_1 , or more generally into S_i when S_i is empty, is performed as follows: If S_{i+1} is empty, we first recursively pull elements into S_{i+1} . If $|S_{i+1}| \leq 2^{i-1}$, we swap the roles of S_{i+1} and the empty S_i , i.e., in constant time move all elements from S_{i+1} to S_i . Otherwise, we select and delete the 2^{i-1} smallest elements from S_{i+1} in time $\mathcal{O}(2^i)$ and let S_i be these (and update p_{i+1} ; here it is crucial that elements are distinct to ensure $\max S_i < p_{i+1} = \min S_{i+1}$).

5.1. Analysis

To prove the amortized performance of this structure we apply the following potential function $\Phi = \varphi^{\text{insert}} + \varphi^{\text{push}} + \varphi^{\text{pull}}$, where

$$\begin{aligned} \varphi^{\text{insert}} &:= \sum_{i=1}^{\ell} \varphi_i^{\text{insert}} & \varphi^{\text{push}} &:= \sum_{i=1}^{\ell} \varphi_i^{\text{push}} & \varphi^{\text{pull}} &:= \sum_{i=1}^{\ell} \varphi_i^{\text{pull}} \\ \varphi_i^{\text{insert}} &:= \max\{0, |S_i| - 5 \cdot 2^{i-1}\} \in [0, 2^{i-1}] & \varphi_i^{\text{push}} &:= \lfloor |S_i|/2^i \rfloor \in [0, 3] \\ \varphi_i^{\text{pull}} &:= \max\left\{0, 2^{i-1} - \sum_{j=1}^i |S_j|\right\} \in [0, 2^{i-1}] \end{aligned}$$

The basic idea is to let each set S_i have a potential. A big set will have a push potential φ_i^{push} that releases one unit of potential when the set is moved to S_{i+1} . An insertion into S_i can trigger a push of S_i . The saved up insertion potential $\varphi_i^{\text{insert}}$ will be released by the push of S_i to cover the increase in pull potential φ_i^{pull} . Finally, pull potential is intuitively associated with each “empty slot” among the first 2^{i-1} “slots” in S_i , with a potential to cover the cost to be recursively filled from each of the subsequent sets S_{i+1}, \dots, S_ℓ . Figure 12 illustrates the relation between set sizes and the different potentials.

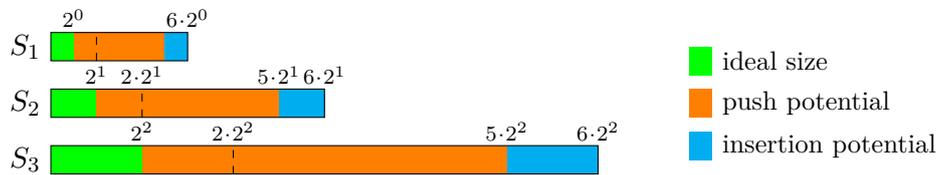


Figure 12: Illustration of set sizes and potentials for the S_i sets.

Lemma 5.1: For a heap with n elements $\ell \leq 1 + \lg n$. ◁

Proof: S_ℓ is created when at least $2^{\ell-1}$ elements are pushed from $S_{\ell-1}$ to S_ℓ during INSERT, i.e., $n \geq 2^{\ell-1}$ implying $\ell \leq 1 + \lg n$. During DELETE-MIN n decreases by one. The concatenation of $S_{\ell-1}$ and S_ℓ when $\ell > 1 + \lg n$, ensures ℓ decreases by one and $\ell \leq 1 + \lg n$ after DELETE-MIN. \square

Lemma 5.2: An INSERT operation, excluding recursive pushing, takes worst-case $\mathcal{O}(\lg \ell)$ time and the change in potential is $\Delta\Phi \leq 2$. \triangleleft

Proof: The binary search for i and appending to S_i takes worst-case $\mathcal{O}(\lg \ell)$ time. Increasing $|S_i|$ by one can at most increase $\varphi_i^{\text{insert}}$ and φ_i^{push} by one, and decrease φ_j^{pull} for each $j \geq i$. It follows that $\Delta\Phi \leq 2$. \square

Lemma 5.3: A DECREASE-KEY operation, excluding recursive pushing, takes worst-case $\mathcal{O}(\lg \ell)$ time and the change in potential is $\Delta\Phi \leq 2$. \triangleleft

Proof: Performing DECREASE-KEY involves two binary searches on the $\ell - 1$ pivots, requiring worst-case $\mathcal{O}(\lg \ell)$ time, and worst-case $\mathcal{O}(1)$ time for moving an element from S_i to S_j , where $1 \leq j \leq i \leq \ell$. The potentials φ_m^{pull} can only change for $j \leq m < i$, in which case they decrease by one, $\varphi_i^{\text{insert}}$ and φ_i^{push} can only decrease, and $\varphi_j^{\text{insert}}$ and φ_j^{push} at most increase by one. It follows that $\Delta\Phi \leq 2$. \square

Lemma 5.4: The recursive pushing of S_i into S_{i+1} , when $|S_i| = 3 \cdot 2^i$, takes worst-case $\mathcal{O}(m - i)$ time and the change in potential is $\Delta\Phi \leq -m + i + 3$, for some m where $i < m \leq \ell + 1$. \triangleleft

Proof: Assuming the recursive pushing continues until pushing elements into S_m for $i < m \leq \ell + 1$, i.e., $|S_j| \geq 2^j$ for $i < j < m$ and $|S_m| < 2^m$ before the push. Each of the $m - i$ pushes takes worst-case constant time to either rename sets or to concatenate two doubly-linked lists.

In the following, recall a bar over a value denotes the value after an operation. We first consider the change to the potential φ^{insert} . Before the push, we have $|S_i| = 6 \cdot 2^{i-1}$, i.e., $\varphi_i^{\text{insert}} = 2^{i-1}$. Since after the push $|\bar{S}_i| = 0$, we have $\bar{\varphi}_i^{\text{insert}} = 0$. For $i + 1 \leq j \leq m$, between 2^{j-1} and $3 \cdot 2^{j-1}$ elements are pushed to S_j . It follows that after the recursive push $|S_j| \leq 5 \cdot 2^{j-1}$, for $i \leq j \leq m$, i.e., $\bar{\varphi}_i^{\text{insert}} = \dots = \bar{\varphi}_m^{\text{insert}} = 0$. It follows that $\Delta\varphi^{\text{insert}} \leq -2^{i-1}$.

For φ^{pull} , note that $|S_j|$ only changes for $i \leq j \leq m$ and only these $\bar{\varphi}_j^{\text{pull}}$ can change. For $i < j < m$, we have $|S_j| \geq 2^{j-1}$ before and after the push, i.e., $\varphi_j^{\text{pull}} = \bar{\varphi}_j^{\text{pull}} = 0$, $|S_m|$ increases, i.e., $\bar{\varphi}_m^{\text{pull}}$ can only decrease, and the maximal increase in φ_i^{pull} is from 0 to 2^{i-1} , i.e., $\varphi^{\text{pull}} \leq 2^{i-1}$.

For φ^{push} , we have $|\bar{S}_i| = 0$, i.e., $\bar{\varphi}_i^{\text{push}} = 0$. For $i < j < m$, since we have $\bar{S}_j = S_{j-1}$ and $|S_{j-1}| \geq 2^{j-1}$ before the recursive push, we have $\varphi_{j-1}^{\text{push}} \geq 1$ and $\bar{\varphi}_j^{\text{push}} \leq \varphi_{j-1}^{\text{push}} - 1$. Since $\bar{\varphi}_i^{\text{push}} = 0$ and $\bar{\varphi}_m^{\text{push}} \leq 2$, it follows that $\Delta\varphi^{\text{push}} \leq 2 - (m - i - 1) = -m + i + 3$. \square

Lemma 5.5: A DELETE-MIN operation, excluding recursive pulling, takes worst-case $\mathcal{O}(1)$ time and the change in potential is $\Delta\Phi \leq \ell$. \triangleleft

Proof: Removing the minimum from S_1 takes worst-case constant time, since $|S_1| \leq 6$. Decreasing $|S_1|$ can only decrease $\varphi_1^{\text{insert}}$ and φ_1^{push} , or leave them unchanged. All other $\varphi_j^{\text{insert}}$ and φ_j^{push} remain unchanged. Decreasing $|S_1|$ by one can at most increase φ_j^{pull} by one for all $1 \leq j \leq \ell$. The total change in potential becomes $\Delta\Phi \leq \ell$.

Since the deletion of an element causes n to decrease by one, we might have $\ell > 1 + \lg n$. In this case DELETE-MIN in worst-case constant time concatenates $S_{\ell-1}$ and S_ℓ , and ℓ decreases by one, ensuring $\ell \leq 1 + \lg n$. Since $\ell > 1 + \lg n$ before the concatenation, we have $|\bar{S}_{\ell-1}| \leq n < 2^{\ell-1}$, and $\varphi_{\ell-1}^{\text{push}} = \varphi_{\ell-1}^{\text{insert}} = 0$ and $\varphi_{\ell-1}^{\text{pull}}$ can only have decreased. It follows that Φ can only decrease by performing this additional check during DELETE-MIN. \square

Lemma 5.6: The recursive pulling of elements into S_1 , when $|S_1| = 0$, takes worst-case $\mathcal{O}(m + 2^{i-1})$ time and the change in potential is $\Delta\Phi \leq -m - 2^{i-1} + 1$, for some i and m , where $1 \leq i < m \leq \ell$. \triangleleft

Proof: Assume when pulling into S_1 we have $|S_1| = \dots = |S_{m-1}| = 0$ and $|S_m| > 0$ for some $1 < m \leq \ell$. Furthermore, let i be the largest $i < m$ where $2^{i-1} < |S_m|$. I.e., S_m will be iteratively swapped with S_{m-1}, \dots, S_{i+1} , and for $1 \leq j \leq i$ we fill S_j by selecting the 2^{j-1} smallest elements from S_{j+1} in $\mathcal{O}(2^j)$ time. It follows that the worst-case time is $\mathcal{O}(m + \sum_{j=1}^i 2^j) = \mathcal{O}(m + 2^i)$.

After the pull, $|S_j| \leq 2^{j-1}$ for $1 \leq j < m$, i.e., $\bar{\varphi}_j^{\text{insert}} = 0$ and $\bar{\varphi}_j^{\text{push}} = 0$. Since $|S_m|$ decreases, we have that $\Delta\varphi_m^{\text{insert}} \leq 0$ and $\Delta\varphi_m^{\text{push}} \leq 0$. In total φ^{insert} and φ^{push} can only decrease.

For φ^{pull} , we have that $\bar{\varphi}_m^{\text{pull}}, \dots, \bar{\varphi}_\ell^{\text{pull}}$ are unchanged, since we only move elements from S_m to S_0, \dots, S_{m-1} . Since at least one element is moved to S_0 , we have φ_j^{pull} decreases by at least one for all $1 \leq j < m$. For $1 \leq j \leq i$, exactly 2^j elements are moved to S_j , and subsequently subsets hereof to S_{j-1}, \dots, S_1 , i.e., we have $\Delta\varphi_j^{\text{pull}} = -2^{j-1}$. It follows that $\Delta\varphi^{\text{pull}} \leq -\sum_{j=1}^i 2^{j-1} - (m-1-i) = -2^i - m + i + 2 \leq -2^{i-1} - m + 1$, which is also a bound on $\Delta\Phi$. \square

By Lemma 5.1 we have $\ell = \mathcal{O}(\lg n)$. Lemmas 5.1, 5.2 and 5.4 imply INSERT takes amortized $\mathcal{O}(\lg \lg n)$ time. Lemmas 5.1, 5.3 and 5.4 imply DECREASE-KEY takes amortized $\mathcal{O}(\lg \lg n)$ time. Finally, Lemmas 5.1, 5.5 and 5.6 imply DELETE-MIN takes amortized $\mathcal{O}(\ell) = \mathcal{O}(\lg n)$ time.

Corollary 5.7: *The amortized time for INSERT and DECREASE-KEY is $\mathcal{O}(\lg \lg n)$ and the amortized time for DELETE-MIN is $\mathcal{O}(\lg n)$.* \triangleleft

6. Conclusion

We presented three priority queue implementation, following a simple partition-based heap framework that uses elementary data structures and algorithms. In particular for the simplest design, the lazy partition heap, we give a fully self-contained and complete description and analysis.

Replacing the binary search on an array by a binary search tree, the implementation works in the pointer-machine model and requires only constant indegree on nodes, and can hence be made persistent efficiently.

Due to its simplicity and linear structure, our implementation appears amenable to adaptations for external memory, but supporting pointer-based decrease-key operations becomes more complicated (as noted by [NPPS11]). We reserve this for future work.

Using our rule as the concatenation rule in quickheaps promises to yield a highly practical and robust priority queue implementation, combining the best of previously known quickheaps variants. A detailed empirical evaluation is planned for future work.

The astute reader may have noticed that INSERT and DECREASE-KEY take amortized constant time, except for the binary searches over the $\mathcal{O}(\lg n)$ pivots. By using a non-pointer-machine model data structure such as a fusion tree [FW93] to store the pivots, the running time can be reduced to $\mathcal{O}(\lg_w \lg n)$, which under the standard assumption that $w = \Omega(\log n)$ results in $\mathcal{O}(1)$ amortized running times for these operations.

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Appendix

A. LP Heap Proof-of-Concept Implementation

To illustrate the simplicity of LP Heaps, we provide a fully self-contained Python implementation below. The code is intended to serve as detailed specification of LP Heaps, tuned for readability. It is not intended to be efficient.

We first show the main code for LP Heaps, the class `LazyHeap`; it uses a doubly-linked list and a selection method which we provide below.

```

1 class LazyHeap:
2     def __init__(self):
3         self.counter = 0
4         self.pivots = []
5         self.sets = [LinkedList()]
6
7     def insert(self, value):
8         node = Node((value, self.counter))
9         self.counter += 1
10        self._insert(node)
11        return node
12
13    def delete(self, node):
14        self._remove(node)
15        self._merge_sets()
16
17    def extract_min(self):
18        S = self.sets[0]
19        if len(S) == 0:
20            raise ValueError("Heap is empty")
21        smallest_node = min(S.iter_nodes(),
22                           key=lambda node: node.value)
23        S.remove(smallest_node)
24        if len(S) >= 2:
25            median, first, second = self._partition_set(S)
26            self.pivots = [median] + self.pivots
27            self.sets = [first, second] + self.sets[1:]
28        self._forget_pivots()
29        return smallest_node.value[0]
30
31    def decrease_key(self, node, new_value):
32        assert new_value <= node.value[0]
33        self._remove(node)
34        node.value = (new_value, node.value[0])
35        self._insert(node)
36
37    def _insert(self, node):
38        S = self._find_set(node.value)
39        S.append(node)
40
41    def _remove(self, node):
42        S = self._find_set(node.value)
43        S.remove(node)
44
```

```

47 def _find_set(self, value):
48     if len(self.sets) == 1 or value < self.pivots[0]:
49         return self.sets[0]
50         low = 0
51     high = len(self.pivots)
52     while low + 1 < high:
53         mid = (low + high) // 2
54         if self.pivots[mid] <= value:
55             low = mid
56         else:
57             high = mid
58     assert value >= self.pivots[low]
59     return self.sets[low + 1]

61 def _partition_set(self, S):
62     median = select(S, -(-len(S) // 2))
63     smaller = []
64     larger = []
65     for node in S.iter_nodes():
66         if node.value < median:
67             smaller.append(node)
68         else:
69             larger.append(node)
70     first = LinkedList()
71     for node in smaller:
72         first.append(node)
73     second = LinkedList()
74     for node in larger:
75         second.append(node)
76     return median, first, second

78 def _forget_pivots(self):
79     new_pivots = []
80     new_sets = [self.sets[0]]
81     outside = 0
82     for pivot, S in zip(self.pivots, self.sets[1:]):
83         if len(new_sets[-1]) == 0 or
84            outside > len(new_sets[-1]) + len(S):
85             new_sets[-1].concat(S) // merge
86         else:
87             outside += len(new_sets[-1])
88             new_pivots.append(pivot)
89             new_sets.append(S)
90     if len(new_sets[-1]) == 0 and len(new_sets) >= 2:
91         new_pivots.pop()
92         new_sets.pop()
93     self.pivots = new_pivots
94     self.sets = new_sets

```

The code above uses a circularly closed doubly-linked list; since the node objects are used as pointers/references to inserted elements, we provide the simple implementation here:

```

1 class Node:
2     def __init__(self, value):
3         self.value = value

```

```

4         self.prev = None
5         self.next = None

7 class LinkedList:
8     def __init__(self):
9         self.head = Node(None)
10        self.tail = Node(None)
11        self.head.next = self.tail
12        self.tail.prev = self.head
13        self.size = 0

15    def __len__(self):
16        return self.size

18    def append(self, node):
19        self.tail.prev.next = node
20        node.prev = self.tail.prev
21        node.next = self.tail
22        self.tail.prev = node
23        self.size += 1

25    def remove(self, node):
26        node.prev.next = node.next
27        node.next.prev = node.prev
28        node.prev = None
29        node.next = None
30        self.size -= 1

32    def concat(self, other):
33        t = self.tail.prev
34        h = other.head.next
35        t.next = h
36        h.prev = t
37        self.tail = other.tail
38        self.size += other.size

40    def iter_nodes(self):
41        at = self.head.next
42        while at is not self.tail:
43            yield at
44            at = at.next

46    def __iter__(self):
47        for node in self.iter_nodes():
48            yield node.value

```

To be self-contained, we also give a simple implementation of the median-of-medians linear-time deterministic selection method [BFP⁺73]. Note that the method is not in-place and should be considered a “pedagogical” implementation. (A practical implementation would certainly want to use a randomized quickselect variant instead.)

```

1 def select(iterable, k):
2     def inner(arr, k):
3         buckets = [sorted(arr[i:i+5]) for i in range(0, len(arr), 5)]

```

```
5     if len(buckets) == 1:
6         return buckets[0][k]

8     medians = [buck[len(buck) // 2] for buck in buckets]
9     pivot = inner(medians, len(medians) // 2)

11    smaller = []
12    larger = []

14    for value in arr:
15        if value < pivot:
16            smaller.append(value)
17        else:
18            larger.append(value)

20    if k < len(smaller):
21        return inner(smaller, k)
22    else:
23        return inner(larger, k - len(smaller))

25    return inner(list(iterable), k)
```