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# Advanced Algorithms: Problems and Solutions

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Mattia Setzu

Giorgio Vinciguerra

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# 1 Range updates

Consider an array  $C$  of  $n$  integers, initially all equal to zero. We want to support the following operations:

- *update*( $i, j, c$ ), where  $0 \leq i \leq j \leq n - 1$  and  $c$  is an integer: it changes  $C$  such that  $C[k] := C[k] + c$  for every  $i \leq k \leq j$ .
- *query*( $i$ ), where  $0 \leq i \leq n - 1$ : it returns the value of  $C[i]$ .
- *sum*( $i, j$ ), where  $0 \leq i \leq j \leq n - 1$ : it returns  $\sum_{k=i}^j C[k]$ .

Design a data structure that uses  $O(n)$  space, takes  $O(n \log n)$  construction time, and implements each operation above in  $O(\log n)$  time. Note that *query*( $i$ ) = *sum*( $i, i$ ) but it helps to reason.

[Hint: For the general case, use the segment tree seen in class, which uses  $O(n \log n)$  space: prove that its space is actually  $O(n)$  when it is employed for this problem.]

## 1.1 First solution

Let  $T$  be a segmented binary tree over a continuous interval  $I : [0, N - 1]$  s.t. its leafs are the points in  $I$ , and the parent of two nodes comprises of their interval:

$$n' \cup n'' = n, n' \cap n'' = \emptyset \text{ s.t. } n \text{ is the parent of } n', n''$$

$T$  will keep track of the prefix sums for every interval. We define a function

$$s' : [0, n - 1] \rightarrow \mathbb{N} \quad (1)$$

that given a node in  $T$  returns the value associated with  $I$ , namely the cumulative sum of that interval.

In order to reduce the computational cost, we introduce a lazy algorithm that doesn't propagate sums over  $T$  as they are streamed in the input, which means  $s'(i)$  might not be accurate at a given time  $t$  for any of the requested operation.

We'll instead either compute over  $T$  or update  $T$  as necessary. Let us define a function to do so:

$$l : \mathbb{N} \rightarrow (\mathbb{N} \cup \{\epsilon\}, \mathbb{N}) \quad (2)$$

to keep track of our lazy sums:

$$s(n) = \begin{cases} \epsilon, - & \text{if no lazy prefix sum is in that interval} \\ k, m & \text{if a lazy sum of } k \text{ is to be propagated to } m \end{cases} \quad (3)$$

The QUERY function is then trivial:

```

1: function QUERY( $I, i, sum$ ):
2:   if  $I.size = 1$  then                                     ▷ Return found value
3:     return  $I.sum$ 
4:   if  $lazy(I), i \in I.left, i \notin I.right$  then           ▷ Lazy on left child
5:      $lazy(I) \leftarrow False$ 
6:     QUERY( $I.left, i, sum + I.sum$ )
7:   if  $lazy(I), i \in I.right, i \notin I.left$  then           ▷ Lazy on right child
8:      $lazy(I) \leftarrow False$ 
9:     QUERY( $I.right, i, sum + I.sum$ )
10:  if  $lazy(I), i \in I.right, i \in I.left$  then               ▷ Lazy on both
11:     $lazy(I) \leftarrow False$ 
12:    QUERY( $I.right, i, j, sum + I.sum$ ) + QUERY( $I.left, i, j, sum + I.sum$ )
13:  if  $!lazy(I), i \in I.left$  then                             ▷ Not lazy on left child
14:    QUERY( $I.left, i, sum$ )
15:  if  $!lazy(I), i \in I.right$  then                           ▷ Not lazy on right child
16:    QUERY( $I.right, i, sum$ )
17:  if  $!lazy(I), i \in I.right, i \in I.left$  then             ▷ Not lazy both

```

```

18:      SUM(I.right, i, sum)
1: function SUM(I, i, j, sum):
2:   if I.size = 1 then                                     ▷ Return
3:     return I.sum + sum
4:   if lazy(I), i ∈ I.left, i ∉ I.right then               ▷ Lazy on left
5:     lazy(I) ← False
6:     SUM(I.left, i, j, sum + I.sum)
7:   if lazy(I), i ∈ I.right, i ∉ I.left then               ▷ Lazy on right
8:     lazy(I) ← False
9:     SUM(I.right, i, j, sum + I.sum)
10:  if lazy(I), i ∈ I.right, i ∈ I.left then               ▷ Lazy on both
11:    lazy(I) ← False
12:    SUM(I.right, i, j, sum + I.sum) + SUM(I.left, i, j, sum + I.sum)
13:  if !lazy(I), i ∈ I.left then                             ▷ Not lazy on both
14:    SUM(I.left, i, sum)
15:  if !lazy(I), i ∈ I.right then                             ▷ Not lazy on both
16:    SUM(I.right, i, sum)
17:  if !lazy(I), i ∈ I.right, i ∈ I.left then             ▷ Not lazy on both
18:    SUM(I.right, i, sum)

1: function UPDATE(I, i, j, k):
2:   if I.size = 1 then                                     ▷ Return
3:     return I.val ← I.val + update
4:     return I.val += update
5:   if lazy(I), i ∈ I.left, i ∉ I.right then               ▷ Lazy on left
6:     lazy(I.left) ← True
7:     I.left.val ← k
8:   if lazy(I), i ∈ I.right, i ∉ I.left then               ▷ Lazy on right
9:     lazy(I.right) ← True
10:    I.right.val ← k
11:  if lazy(I), i ∈ I.right, i ∈ I.left then               ▷ Lazy on both
12:    lazy(I) ← True
13:    I.val ← k
14:  if !lazy(I), i ∈ I.left then                             ▷ Not lazy
15:    UPDATE(I.left, i, update)
16:    update(I.left, i, update)
17:  if !lazy(I), i ∈ I.right then                             ▷ Not lazy
18:    UPDATE(I.right, i, update)
19:    update(I.right, i, update)
20:  if !lazy(I), i ∈ I.right, i ∈ I.left then             ▷ Not lazy
21:    UPDATE(I.right, i, update)
22:    update(I.right, i, update)

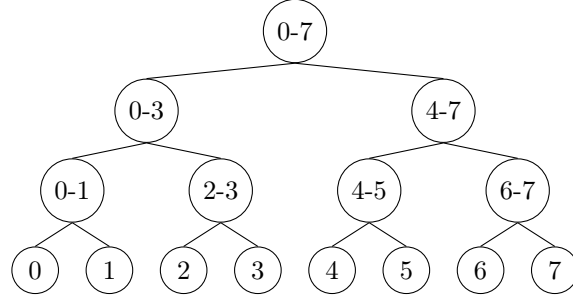
```

## 1.2 Second solution

We use a segment binary tree  $T_I$  over the interval  $I = [0, n-1]$ , that is, a tree whose leaves are the points in  $I$  and the parent of two nodes is the union of their interval. More formally, if  $x.interval$  denotes the attribute *interval* of the node  $x$ ,

1.  $l.interval \cup r.interval = n.interval \iff n$  is the parent of  $l$  and  $r$ ;
2. if  $x$  and  $y$  are leaves, then  $x.interval \cap y.interval = \emptyset \wedge |x.interval| = |y.interval| = 1$ .

For example, the segment tree for  $I = [0, 7]$  is the following:



We associate with each node  $x$  of  $T_I$  some attributes:  $x.sum$ , that stores  $\sum_{i \in x.interval} C[i]$ ; and  $x.lazy$ , that stores a value that need to be propagated to each descendant of  $x$ . This means that  $x.sum$  might not be accurate at a given time for any of the requested operation.

**Range operations.** In both operations we traverse the tree recursively starting from the root and, at each recursive step on any internal node  $x$ , if  $x.lazy \neq 0$ : we set  $x.sum \leftarrow x.sum + |x.interval| \times x.lazy$ , we propagate the lazy information to  $x$ 's children  $x.left.lazy \leftarrow x.lazy$ ,  $x.right.lazy \leftarrow x.lazy$  and, finally, we reset the information  $x.lazy \leftarrow 0$ . Afterwards, if the operation is

- $sum(i, j)$ , we do the following:
  1. if  $x.interval \cap [i, j] = \emptyset$ , we return 0;
  2. if  $x.interval \subseteq [i, j]$ , we return  $x.sum$ ;
  3. otherwise we repeat the procedure  $sum(i, j)$  on  $x.left$  and  $x.right$ , returning the sum of these calls.
- $update(i, j, c)$ , we do the following:
  1. if  $x.interval \cap [i, j] = \emptyset$ , we stop the recursion on this subtree;
  2. if  $x.interval \subseteq [i, j]$ , we update  $x.sum \leftarrow |x.interval| \times c$ , and  $x.left.lazy \leftarrow x.left.lazy + c$ ,  $x.right.lazy \leftarrow x.right.lazy + c$
  3. otherwise we repeat the procedure  $update(i, j, c)$  on  $x.left$  and  $x.right$ .

The space occupied by  $T_I$  is  $\sum_{i=0}^{\log_2 n} n/2^{-i} = 2n - 1$ .

## 2 Depth of a node in a random search tree

A random search tree for a set  $S$  can be defined as follows: if  $S$  is empty, then the null tree is a random search tree; otherwise, choose uniformly at random a key  $k$  as root, and the random search trees on  $L = \{x \in S : x < k\}$  and  $R = \{x \in S : x > k\}$  become, respectively, the left and right subtree of the root  $k$ . Consider the randomized QuickSort discussed in class and analyzed with indicator variables CLRS 7.3, and observe that the random selection of the pivots follows the above process, with indicator variables, prove that:

1. the expected depth of a node (i.e. the random variable representing the distance of the node from the root) is nearly  $2 \ln n$ ;
2. the expected size of its subtree is nearly  $2 \ln n$  too, observing that it is a simple variation of the previous analysis;
3. the that the probability that the depth of a node exceeds  $c 2 \ln n$  is small for any given constant  $c > 1$ .

### 2.1 Proof with indicator variable

**Prove that the expected depth of a node is nearly  $2 \ln n$ .**

*Proof.* Let  $z_m$  the  $m$ th smallest element in  $S$  and

$$X_{ij} = \begin{cases} 1 & \text{if } z_j \text{ is an ancestor of } z_i \text{ in the random search tree} \\ 0 & \text{otherwise} \end{cases}$$

The depth of the node  $i$  in the tree is given by the number of its ancestors:

$$X = \sum_{\substack{j=1 \\ j \neq i}}^n X_{ij} \quad (4)$$

Note that the depth of a node is also equal to the number of comparison it's involved in (in other words, the number of times it became the left or the right child of a randomly chosen pivot).

Once a pivot  $k$  is chosen from  $S$ ,  $S$  is partitioned in two subsets  $L$  and  $R$ . The elements in the set  $L$  will not be compared with the elements in  $R$  at any subsequent time. The event  $E_1 = "z_j \text{ is an ancestor of } z_i \text{ in the random search tree}"$  occurs if  $z_j$  and  $z_i$  belongs to the same partition *and*  $z_j$  was chosen as pivot before  $z_i$ . The probability that  $E_1$  occurs, since it is the intersection of two events, can be upper bounded by:

$$\Pr\{z_j \text{ was chosen as pivot before } z_i\} = \frac{1}{\text{size of the partition}} \leq \frac{1}{|j - i| + 1}$$

because pivots are chosen randomly and independently, and because the partition that contains both  $z_j$  and  $z_i$  must contain *at least* the  $|j - i| + 1$  numbers between  $z_j$  and  $z_i$ .

Taking expectations of both sides of (4), and then using linearity of expectation, we have:

$$\begin{aligned} \mathbb{E}[X] &= \sum_{\substack{j=1 \\ j \neq i}}^n \mathbb{E}[X_{ij}] \\ &= \sum_{\substack{j=1 \\ j \neq i}}^n \Pr\{z_j \text{ is an ancestor of } z_i \text{ in the random search tree}\} \\ &\leq \sum_{\substack{j=1 \\ j \neq i}}^n \frac{1}{|j - i|} \\ &= \sum_{j=1}^{i-1} \frac{1}{i - j} + \sum_{j=i+1}^n \frac{1}{j - i} \end{aligned}$$

With the change of variables  $l = i - j$  and  $m = j - i$ :

$$= \sum_{l=1}^{i-1} \frac{1}{l} + \sum_{m=1}^n \frac{1}{m} \approx 2 \ln n$$

□

**Prove that the expected size of its subtree is nearly  $2 \ln n$  too, observing that it is a simple variation of the previous analysis.**

*Proof.* The size of the subtree of a randomly chosen pivot of  $z_j \in S$  is given by the number of its descendants. Since (4) is the number of ancestors of a node  $z_i$ , we can find the number of descendants of  $z_j$  by changing the summation from  $j = 1, \dots, n$  to  $i = 1, \dots, n$ . □

**Prove that the probability that the depth of a node exceeds  $c \ln n$  is small for any given constant  $c > 1$ .**

*Proof.* We apply the Theorem 1 below with  $(1 + \delta) = c$  and  $\mu = 2 \ln n$ :

$$\mathbb{P}[X > 2c \ln n] < \left( \frac{e^{c-1}}{c^c} \right)^{2 \ln n}$$

Since  $\lim_{n \rightarrow +\infty} a^n = 0$  with  $0 < a < 1$ , and  $\lim_{n \rightarrow +\infty} 2 \ln n = +\infty$ , we have to prove that, for any  $c > 1$

$$0 < \frac{e^{c-1}}{c^c} < 1$$

The first inequality is simple to verify, for the latter observe that

$$\frac{e^{c-1}}{c^c} < 1 \iff e^{c-1} < c^c \iff c - 1 < c \ln c \iff c(1 - \ln c) < 1.$$

□

**Theorem 1** (Chernoff Bound). *Let  $X_1, \dots, X_n$  be independent Poisson trials such that, for  $1 \leq i \leq n$ ,  $\mathbb{P}[X_i = 1] = p_i$ , where  $0 < p_i < 1$ . Then, for  $X = \sum_{i=1}^n X_i$ ,  $\mu = \mathbb{E}[X] = \sum_{i=1}^n p_i$  and any  $\delta > 0$ ,*

$$\mathbb{P}[X > (1 + \delta)\mu] < \left( \frac{e^\delta}{(1 + \delta)^{(1 + \delta)}} \right)^\mu.$$

## 2.2 Recursive balanced proof

Let  $n$  be the number of nodes in the input list  $l$ ,  $h = \log_2(n)$  the height of a balanced tree over  $l$ ,  $T(p)$  the tree built over the permutation  $p$  of pivots,  $d(m)$  be the positional distance of a value  $m$  of a partition from the median value of the said partition. Then the following holds:

- $height(T) = h \iff |T.left| = |T.right| \pm 1$  Trivially, let  $r$  be the root of a 3-nodes partition: then, if the partition is unbalanced, the lesser one will comprise of 0 nodes, while the greater one of 2, which implies that  $height(T.right) = 2$ .
- $P = \text{pivot}, d(m) = \pm k \implies height(T.left) = height(T.right) \pm k$ . Recursively from the previous statement, a partition unbalanced of one element generates subtrees whose levels differ on a factor of 1. By iterating recursively, their subtrees, if unbalanced by 1, will yield one more level difference. Over  $k$  unbalanced pivots on a single subtree, at most  $k$  levels will be added to  $h$ .
- By the previous statement, it follows that  $\nexists T, T' : height(T) > height(T'), T \text{ balanced}, T' \text{ unbalanced}$ . As stated, let  $T', T$  be the unbalanced/balanced tree respectively; let us cheat with  $T$  and switch the root pivot with the first element in its subtree. Now, let us prove by contradiction that  $T$  can't stay balanced and that its height will increase. By shifting the tree to the left we have deprived  $T.right$  of either 0 levels (in case  $T.right$  is able to switch every pivot in its tree with its right subtree root, ending with the rightmost leaf in its subtree) or 1, in case no rightmost leaf is present. Therefore  $height(T) < height(T')$ .



- The completely unbalanced tree is the tree with the most levels. By taking partitions of size 0 we constantly force, at each level, one subtree to disappear. Therefore, its level(s) has to be necessarily transferred to its brother. We then have exactly one node per level, therefore  $n$  levels.

**Behaviour on random permutations** Now let us analyse how the tree depth varies according to random pivot selection. We start by applying the 2.2k-distance to a tree  $T$  with  $n = 3$  nodes. Trivially,  $\text{height}(T)$  with balanced tree is equal to two. Now, let us pick either the lowest or the greatest pivot possible: the tree is unbalanced towards either the left or the right, but  $\text{height}(T) = 2$  in both cases. As the reader can see from 2.2, the distance works in absolute value; it is then clear how, at every permutation for a pivot  $p$ , out of the  $n$ , there are 2 that generate a tree of the same height:  $p = d(P) + k, p = d(P) - k$ . Given that at every iteration a node  $x$  in a completely unbalanced tree  $T'$  has a probability of  $\frac{1}{n-i}$ , we can define the probability of  $x$  being a pivot at level  $l$  as:

$$P(x_k) = \frac{1}{n-l} \quad (5)$$

Now, in order for  $x$  not to be chosen as pivot in the previous  $l-1$  levels we have:

$$P(x_k) = \sum_{k=1}^{l-1} \left( \frac{1}{n-l+1} \right) \quad (6)$$

Given the height of  $T$ , the (harmonic) partial series converges to  $\ln(n) + 1$ . Let us now add a root  $r$  s.t.  $T'.\text{right} = T, T'.\text{left} = T$ . We now have to consider the mirror case  $\ln(n') + \ln(n')$ , given by the previous  $n' = n/2$  in the logarithm, since the number of nodes doubled, the  $+1$  removed for both, since now neither of  $T'.\text{left}, T'.\text{right}$  is the root, and a  $+1$  added since a new level has been added.

**Upper bound.** By hypothesis,

$$\mathbb{E}[d(x) > 2c \ln(n)] \ll 1 \quad (7)$$

By definition the ancestor of a node  $i$  are independent random variables, and we can apply the Chernoff bounds over the set  $x : d(x) \geq 2 \ln(n)$  of random variables determining the expected distance of nodes.

$$\mathbb{P}[X \geq c \mathbb{E}[X]] < e^{-c \ln(\frac{c}{e}) \mathbb{E}[X]}$$

Let us consider  $X = 1 \forall i == \ln(n)$ , the expected depth of  $\ln(n)$ , then

$$\mathbb{P}[X \geq c \ln(n)] < e^{-c \ln(\frac{c}{e}) \ln(n)}$$

### 3 Karp-Rabin fingerprinting on strings

Given a string  $S \equiv S[0 \dots n-1]$ , and two positions  $0 \leq i < j \leq n-1$ , the longest common extension  $lce_S(i, j)$  is the length of the maximal run of matching characters from those positions, namely: if  $S[i] \neq S[j]$  then  $lce_S(i, j) = 0$ ; otherwise,  $lce_S(i, j) = \max\{l \geq 1 : S[i \dots i+l-1] = S[j \dots j+l-1]\}$ . For example, if  $S = \text{abracadabra}$ , then  $lce_S(1, 2) = 0$ ,  $lce_S(0, 3) = 1$ , and  $lce_S(0, 7) = 4$ . Given  $S$  in advance for preprocessing, build a data structure for  $S$  based on the Karp-Rabin fingerprinting, in  $O(n \log n)$  time, so that it supports subsequent online queries of the following two types:

- $lce_S(i, j)$ : it computes the longest common extension at positions  $i$  and  $j$  in  $O(\log n)$  time.
- $equal_S(i, j, l)$ : it checks if  $S[i \dots i+l-1] = S[j \dots j+l-1]$  in constant time.

Analyze the cost and the error probability. The space occupied by the data structure can be  $O(n \log n)$  but it is possible to use  $O(n)$  space. [Note: in this exercise, a one-time preprocessing is performed, and then many online queries are to be answered on the fly]

#### 3.1 First solution

In order to save computational cycles on checks over ranges we use a similar structure to the one in the range updates: we compute the hashing on the first character in  $O(1)$  time, then roll the hash through the  $n-1$  remaining characters through  $nO(1)$  operations. We call  $H$  this array; we also denote  $h_k$  as the function  $ca^i$  computing the Rabin-Karph hash of a string  $s$ . The reader shall now see that  $\exists h^{-1}(s)$ : that is,  $h$  is invertible in  $O(1)$ . The entries  $h[i] = \sum_{i \in [0, n-1]} (h(i))$  have cumulative hash and the following properties hold:

- $h[s[i]] = (h[i] - h[i-1])/a^{-1}, a^{-1} = a^1$
- $h[i..j] - h[k..l] = (h[l] - h[k-1])/a^{-1} - (h[j] - h[i-1])/a^{-1}, a^{-1} = \text{modular inverse}$

EQUALS works on cumulative hashes, subtracting them and scaling them accordingly, as our *rabin* function multiplies by an  $a^i$  constant.

```

1: function EQUALS( $i, j, length$ ):
2:    $h_i = h[i + length] - h[i - 1]$ 
3:    $h_j = h[j + length] - h[j - 1]$ 
4:    $h^i = h_i / inv(a, i, l)$ 
5:    $h^j = h_j / inv(a, j, l)$ 
   return  $h^i - h^j == 0$ 
6: function INV( $h, k, l$ ): return  $h^{k-l}$ 
```

LCE works on cumulative hashes, checks the equality on the middle element of the strings and runs recursively on the half with different hashing. We define LCE as an auxiliary function

```

1: function LCE( $i, j, l$ ):
2:    $eq = \text{EQUALS}(i, j)$ 
3:   if  $eq$  then return  $l$ 
4:   else if  $\neg \text{EQUALS}((j-i)/2, (n-j)/2, l)$  then return  $\text{EQUALS}((j-i)/2, (n-j)/2)$ 
5:   else  $(j-i)/2 + \text{return EQUALS}((j-i)/2, (n-j)/2)$ 
6:    $h_i = h[i + length] - h[i - 1]$ 
7:    $h_j = h[j + length] - h[j - 1]$ 
8:    $h^i = h_i / inv(a, i, l)$ 
9:    $h^j = h_j / inv(a, j, l)$ 
   return  $h^i - h^j == 0$ 
10: function INV( $h, k, l$ ): return  $h^{k-l}$ 
```

#### 3.2 Second solution

We interpret the string  $S \in \Sigma^*$  as the polynomial  $S(z) = \sum_{i=0}^{n-1} S[i]z^i$ , where  $S[i]$  is a symbol encoded with a number in  $\{0, \dots, |\Sigma| - 1\}$ . Given a prime number  $p$  greater than both  $2n$  and  $|\Sigma|$ , we define the

fingerprint on strings  $h_p$  as:

$$h_p(S) = h_p \left( \sum_{i=0}^{n-1} S[i] \cdot z^i \right) = \left( \sum_{i=0}^{n-1} S[i] \cdot z^i \right) \bmod p$$

where  $z$  is randomly chosen from  $\mathbb{Z}_p$ .

If two strings  $A$  and  $B$  are equal, then also their fingerprints are equal. If they are different, the probability that their fingerprints are the same is at most  $\frac{n}{p} < \frac{n}{2n} = \frac{1}{2}$ , because there are  $p$  choices for  $z$  and at most  $n$  roots of the polynomial  $P(z) = A(z) - B(z)$  (follows from the fundamental theorem of algebra and the fact that  $A(z)$  and  $B(z)$  have both degree  $n$ ) [Motwani10].

**Queries.** To implement efficiently the queries on  $S$ , we use an array  $H$  with the hashes of the string's prefixes, formally:

$$H[i] = h_p(S[0..i]) \quad \forall i \in [0, n-1]$$

where  $S[i..j]$  denotes the substring  $S[i]S[i+1]\dots S[j]$ .

With the array  $H$  we can compute in constant time the fingerprint of any substring of  $S$ , hence we can decide — with a probability of error —  $\text{equals}(i, j, l)$ , by comparing the substrings' fingerprints. To compute  $\text{lce}_S(i, j)$ , we search with  $O(\log n)$  different values of  $l$  inside the interval  $[0, \min(j-i, n-j)]$  until the result is found. Listing 1 shows an implementation in Python.

**Analysis.**  $H$  is built in  $O(n)$  time, because at each iteration the fingerprint in  $H[i]$  is computed in constant time from the fingerprint in  $H[i-1]$ . The space needed by the solution is  $O(n \log p)$ , because we need to store the array  $H$ , that has  $n$  entries of size  $\log_2 p$ .

Listing 1: Source code to compute the longest common extension efficiently.

```

1 p = 179426549
2 z = 3
3 s = "abracadabra"
4
5 def h(s):
6     return sum(ord(c) * z ** i for i, c in enumerate(s)) % p
7
8 H = [ord(s[0]) % p] # Array with hashes of prefixes
9 exp = z
10 for i in range(1, len(s)): # O(n)
11     H.append((H[i-1] + (ord(s[i]) * exp) % p))
12     exp = exp * z
13
14 def rollLeft(j, count): # Compute the hash of the string s[j:j+count]
15     assert count > 0 and j + count <= len(s)
16     if j == 0:
17         return H[count-1]
18     return (H[j+count-1] - H[j-1]) / z ** j
19
20 def equal(i, j, l): # O(1)
21     assert i < j and j + l - 1 < len(s) and l > 0
22     hs1 = rollLeft(i, l) # hash of s[i:i+l]
23     hs2 = rollLeft(j, l) # hash of s[j:j+l]
24     return hs1 == hs2
25
26 def lceaux(i, j, intsize, maxlce): # O(log n)
27     if (intsize <= 1):
28         return 1 if equal(i, j, 1) else 0
29     if equal(i, j, intsize): # Check if there is a longer common extension
30         if (intsize + intsize / 2 >= maxlce):

```

```

31         return maxlce
32     return lceaux(i, j, intsize + intsize / 2, maxlce)
33 else: # Check if there is a shorter common extension
34     return lceaux(i, j, intsize - intsize / 2, intsize)
35
36 def lce(i, j):
37     maxlce = min(j - i, len(s) - j)
38     if (equal(i, j, maxlce)):
39         return maxlce
40     return lceaux(i, j, maxlce, maxlce)

```

## 4 Hashing sets

Your company has a database  $S \subseteq U$  of keys. For this database, it uses a hash function  $h$  uniformly chosen at random from a universal family  $\mathcal{H}$  (as seen in class); it also keeps a bit vector  $B_S$  of  $m$  entries, initialized to zeroes, which are then set  $B_S[h(k)] = 1$  for every  $k \in S$  (note that collisions may happen). Unfortunately, the database  $S$  has been lost, thus only  $B_S$  and  $h$  are known, and the rest is no more accessible. Now, given  $k \in U$ , how can you establish if  $k$  was in  $S$  or not? What is the probability of error? [Note: you are not choosing  $k$  and  $S$  randomly as they are both given... randomization here is in the choice of  $h \in \mathcal{H}$  performed when building  $B_S$ .] Under the hypothesis that  $m \geq c|S|$  for some  $c > 1$ , find the expected number of 1s in  $B_S$  under a uniform choice at random of  $h \in \mathcal{H}$ .

**Solution.** Trivially for  $B_S[h(k)] = 0$  we can answer FALSE with  $\mathbb{P}[\text{error}] = 0$ . Let us analyse the opposite case,  $B_S[h(k)] = 1$ . Let  $i \in [0, m]$  be some index s.t.  $B_S[i] = 1$ , and let  $cl_S(i)$  be the list of  $k \in S : h(k) = i$  for some set  $S \in \mathcal{P}(S)$ . We can then denote the sets  $cl_U := \{k_U : k_U \in U, h(k) = i, cl_S := \{k_S : k_S \in S, h(k) = i\}$ ; it is trivial to show that

- $cl_S \subseteq cl_U$  as  $S \in U$ .
- $|cl_S(k)| \leq |cl_U(k)| \forall k \in U$  as  $S \in U$ .

Let us not try and estimate  $|cl_S(k)|$ : given  $h$  is universal, we have an expected value of collisions of  $\mathbb{E}[X_k] \approx \frac{1}{m} \forall k \in S$ , that is

$$\begin{aligned}\mathbb{P}[h(k^0) = c] &= \frac{1}{m} \\ \mathbb{P}[h(k^1) = c] &= \frac{1}{m^2} \\ \mathbb{P}\left[h(k^{i-1}) = c, h(k^i) = c\right] &= \frac{1}{m^i} \\ \mathbb{P}[h(k) = c, \forall k \in S] &= \frac{1}{m^{|S|}}\end{aligned}$$

We can similarly compute the probability of not collision by simply replacing  $\frac{1}{m}$  with  $(1 - \frac{1}{m})$ :  $\mathbb{P}[h(k) \neq c, \forall k \in S] = 1 - \frac{1}{m^{|S|}}$ .

Given our estimate of the collision list, we can now compute an estimate of the error probability: by 4, we give an erroneous answer whenever  $k \in cl_U(k), k \notin cl_S(k)$ , that is we have a margin of error of  $cl_U(k) \setminus cl_S(k)$  whose size is  $|cl_U(k)| - |cl_S(k)|$ . Given the set of  $k$  for which  $h(k) = i$  the *bad answers* are then

$$1 - \mathbb{P}[\text{good answer}] = 1 - \left(1 - \frac{1}{m}\right)^{|S|} \quad (8)$$

We now provide a lower and upper bound for the said value. The lower bound is given by the *perfect hash* with no collisions: given  $e$  = number of 1 in  $B_S$ ,  $e$  is the lowerbound. Provided that  $m \geq c|S|$  for some  $c > 1, |S| \leq \frac{m}{c}$ :

$$\mathbb{P}[\text{collision over } i] = \alpha_S = \frac{\frac{m}{c}}{m} = \frac{1}{c} \quad (9)$$

Therefore

$$e \leq 1 - \frac{1}{m^{|S|}} \leq \frac{1}{c} \quad (10)$$

**Expected number of 1 in  $B_S$ .** We define a random variable  $X_k$ :

$$X_k = \begin{cases} 1 & \text{if } B_S[k] = 1 \\ 0 & \text{otherwise} \end{cases}$$

and one over the ones in  $B_S$ ,  $X$ :

$$X = \sum_{k=0}^{m-1} X_k \quad (11)$$

We compute the relative expected value  $\mathbb{E}[X]$ :

$$\begin{aligned}\mathbb{E}[X] &= \mathbb{E}\left[\sum_{k=0}^{m-1} X_k\right] \\ &= \sum_{k=0}^{m-1} \mathbb{P}[B_s[k] = 1] \\ &= \sum_{k=0}^{m-1} \frac{\alpha}{m} \leq \frac{m}{c}\end{aligned}$$

where  $\mathbb{P}[B_s[k] = 1] = \frac{\alpha}{m}$  as  $\alpha$  are the favorable cases (i.e. the collisions for a generic bucket  $B_s[i]$ ), and  $m$  is the size of  $B_s$ .

## 5 Family of uniform hash functions

The notion of pairwise independence says that, for any  $x_1 \neq x_2, c_1, c_2 \in \mathbb{Z}_p$ , we have that

$$\mathbb{P}[h(x_1) = c_1, h(x_2) = c_2] = \mathbb{P}[h(x_1) = c_1] \times \mathbb{P}[h(x_2) = c_2] \quad (12)$$

In other words, the joint probability is the product of the two individual probabilities. Show that the family of hash functions  $\mathcal{H} = \{h_{a,b}(x) = ((ax+b) \bmod p) \bmod m : a \in \mathbb{Z}_p^*, b \in \mathbb{Z}_p\}$  is *pairwise independent* where  $p$  is a sufficiently large prime number ( $m+1 \leq p \leq 2m$ ).

### 5.1 First proof

By linear algebra,  $a + (kp) \bmod p = c \forall k \in \mathbb{N}$ . If we were to cap  $a + pk \bmod p = c \forall k \in K, K_N = k_i : k_i < N$

$$\mathbb{P}[h(x_i) = c_i] = \frac{1}{m^2} = \mathbb{P}[h(x_1) = c_1, h(x_2) = c_2]$$

Given  $x_i, d$ , we define as  $m_i = (ax_i + b)$ ; since  $(ax_i + b)$  is a *linear transformation*  $m_i$  is unique. It follows trivially that there are  $\frac{p}{m}$  values for which  $m_i \bmod p = d$  since  $p > m$  and by 5.1. The same goes for  $x_1, x_2$ :

$$\begin{aligned} (ax_1 + b) \bmod p &= d \\ (ax_2 + b) \bmod p &= e \end{aligned}$$

By the Chinese remainder theorem the above system has only one solution for the variable  $(a, b)$  over the  $(p)(p-1)$  possible pairs, therefore  $(ax_1 + b) \bmod p = d$  and  $(ax_2 + b) \bmod p = e$  for  $\approx \frac{p}{m}$  cases each. Since  $d, e$  are independent

$$((ax_2 + b) \bmod p = d) \wedge ((ax_2 + b) \bmod p = e) \text{ for } \frac{p}{m}, \frac{p}{m} = \frac{p^2}{m^2}$$

Over all the possible choices of  $(a, b)$ :

$$\mathbb{P}[(ax_2 + b) \bmod p = d \wedge (ax_2 + b) \bmod p = e] = \frac{\frac{p^2}{m^2}}{p(p-1)}$$

We now prove  $\mathbb{P}[h(x_1) = c_1] = \frac{1}{m}$ . Of all the possible  $m$  buckets, one and only one is the one we look for:  $\mathbb{P}[l \bmod m = c_1] = \frac{1}{m}$ . As we've seen by 5.1 at most  $\frac{p}{m}$  such  $l$  exist for  $(ax + b) \bmod p$ , therefore  $\mathbb{P}[h(x_i) = c_i] = \frac{\frac{p}{m}}{m} \approx \frac{1}{m}$ . Which computes  $\approx \frac{1}{m^2}$  and proves the assumption under the said approximation.

### 5.2 Second proof

*Proof.* We will show that both sides of (12) are approximately equal to  $\frac{1}{m^2}$ .

Given a hash function  $h_{a,b} \in \mathcal{H}$ , and two distinct inputs  $x_1, x_2 \in \mathbb{N}$ , let:

$$\begin{aligned} r &= (ax_1 + b) \bmod p \\ s &= (ax_2 + b) \bmod p \end{aligned}$$

Notice that  $r \neq s$  because  $r - s \equiv a(x_1 - x_2) \pmod{p}$ , both  $a$  and  $x_1 - x_2$  are nonzero modulo a prime number, and so their product is nonzero.

Since there are a total of  $p(p-1)$  pairs of  $(r, s)$  such that  $r \neq s$  ( $p^2$  choices subtracted the number of pairs where  $r = s$ ), there is a one-to-one correspondence between pairs  $(a, b) \in \mathbb{Z}_p^* \times \mathbb{Z}_p$  and pairs  $(r, s)$ . Therefore, for the left-hand side of (12):

$$\mathbb{P}_{h_{a,b} \in \mathcal{H}}[(h(x_1) = c_1, h(x_2) = c_2)] = \mathbb{P}_{r \neq s \in \mathbb{Z}_p^2}[r \bmod m = c_1, s \bmod m = c_2]$$

There are about  $p/m$  values of  $r$  that satisfy  $r \bmod m = c_1$ , and the same goes for  $c_2$  and  $s$ . Hence, there are about  $(p/m)^2$  choices for the pair  $(r, s) \in \mathbb{Z}_p^2$  that satisfy  $r \bmod m = c_1 \wedge s \bmod m = c_2$ . Dividing the favorable cases with the all possible cases:

$$\frac{\lfloor p/m \rfloor^2}{p(p-1)} \leq \mathbb{P}[r \bmod m = c_1, s \bmod m = c_2] \leq \frac{(\lfloor p/m \rfloor + 1)^2}{p(p-1)}$$

Thus,  $\mathbb{P}[h(x_1) = c_1, h(x_2) = c_2] \approx \frac{1}{m^2}$ .

The right-hand side of (12) is equal to  $\frac{1}{m^2}$ , because both probabilities are  $\frac{1}{m}$ . Take for example  $\mathbb{P}[h(x_2) = c_2]$ : there are about  $p/m$  values of  $s$  that satisfy  $s \bmod m = c_2$ , out of  $p$  values for  $s$ . Hence:

$$\mathbb{P}[h(x_2) = c_2] = \mathbb{P}[s \bmod m = c_2] \approx \frac{\frac{p}{m}}{p} = \frac{1}{m}$$

□



## 6 Deterministic data streaming

Consider a stream of  $n$  items, where items can appear more than once. The problem is to find the most frequently appearing item in the stream (where ties are broken arbitrarily if more than one item satisfies the latter). For any fixed integer  $k \geq 1$ , suppose that only  $k$  items and counters can be stored, one item per memory cell, where each counter can use only  $O(\text{polylog}(n))$  bits (i.e.  $O(\log^c n)$  for any fixed constant  $c > 0$ ): in other words, only  $b = O(k \cdot \text{polylog}(n))$  bits of space are available. (Note that, even though we call them counters, they can actually contain any kind of information as long as it does not exceed that amount of bits.) Show that the problem cannot be solved deterministically under the following rules: the algorithm can only use  $b$  bits, and read the next item of the stream, one item at a time. You, the adversary, have access to all the stream, and the content of the  $b$  bits stored by the algorithm: you cannot change those  $b$  bits and the past, namely, the items already read by the algorithm, but you can change the future, namely, the next item to be read. Since the algorithm must be correct for any input, you can use any amount of streams to be fed to the algorithm and as many distinct items as you want. [Hint: it is an adversarial argument based on the fact that, for many streams, there can be a tie on the items.]

**Solution.** We, the adversaries, can decide the alphabet  $\Sigma$  and the content of the stream  $s \in \Sigma^*$ . The idea is to create a stream generator that forces any deterministic algorithm to reach a configuration with at least two different streams. In fact, the number of configurations of memory for this class of algorithms is bounded because of  $b$  and, by the pigeonhole principle, there exist at least two streams that cause the algorithm to transition the same configuration.

Since there are  $2^b$  possible configurations of the memory, we choose to create a class of  $2^b$  streams of length  $n$ . Each stream  $s$  has two occurrences of any symbol chosen from an alphabet  $\Sigma_s$  of size  $n/2$ . Therefore, every stream has a tie the frequencies of its elements (until we generate the next one). The alphabets  $\Sigma_1, \Sigma_2, \dots, \Sigma_{2^b}$  are all subset of  $\Sigma$ , that satisfies the following condition:

$$\binom{|\Sigma|}{n/2} > 2^b$$

With this condition we can generate  $2^b$  streams (from just as many alphabets) such that, for any pair of streams, there exist at least one symbols that appear in one stream but not in the other.

We can apply the pigeonhole principle: suppose that the same configuration is reached in two different executions of the algorithm on two streams with symbols, respectively, from  $\Sigma_i$  and  $\Sigma_j$ . When we will emit as  $(n+1)$ -th character of the stream the symbol  $a \in \Sigma_i$  such that  $a \notin \Sigma_j$ , and query the algorithm for the most frequent item, we will receive an answer  $\mathcal{A}$ . The algorithm is deterministic, therefore in the same configuration the answer will be the same for each stream. If  $\mathcal{A} \neq a$ , then the algorithm gave the wrong answer for the first stream. Otherwise, if  $\mathcal{A} = a$ , then it's correct for the first stream, but not for the second one, because  $a \notin \Sigma_j$ .

stream with symbols from $\Sigma_i$	...	d	d	e	e	f	f	g	g	h	h
stream with symbols from $\Sigma_j$	...	4	4	5	5	6	6	7	7	8	8

## 7 Special case of most frequent item in the stream

Suppose to have a stream of  $n$  items, so that one of them occurs  $> \frac{n}{2}$  times in the stream. Also, the main memory is limited to keeping just two items and their counters, plus the knowledge of the value of  $n$  beforehand. Show how to find deterministically the most frequent item in this scenario.

Hint: since the problem cannot be solved deterministically if the most frequent item occurs  $\leq \frac{n}{2}$  times, the fact that the frequency is  $> \frac{n}{2}$  should be exploited.

**Solution.** We'll use the notation of the previous section. Before illustrating our solution we prove that the following hold:

- **Given the  $j^1$  element,  $C(j^1) > C(j^i) \forall i \in S$ .**
- Following the previous,  $C(j^i) < C(j^i), i > 1$  element.
- **There are at most  $\frac{n}{2} - 1$  elements in  $S$ :** given that  $j^1$  appears at least  $\frac{n}{2} + 1$  times, in the best case scenario every other  $\frac{n}{2} - 1$  element is different, giving  $\frac{n}{2} - 1$  different elements.
- **$j^1$  has sub-stream dominance:**  $\exists S' \subseteq S : j^1$  is the most frequent item  $\in S'$ : trivially, since it is the most frequent element,  $S'$  is always a sub-stream dominant sub-stream.

**Recursive sub-streams** It is trivial to show that if we were to have  $\frac{n}{2}$  counters,  $\sum_{i=0, i \neq 1}^{\frac{n}{2}} (C(j^i) < C(j^1))$  by 7. Though trivial, by combining it with 7 we can assert that given  $S' = S[0, \frac{n}{2}]$ ,  $S'' = S[\frac{n}{2}, n]$ ,  $k =$  number of elements  $\in S$

- $j^1$  is the sub-stream dominant element in either  $S', S''$ : by contradiction, let us assume that is false. Then  $\exists j_o \neq j^1 : C(j_o) \text{ in } S' > C(j^1), \exists j_o : C(j_o) \text{ in } S'' > C(j^1)$ ; given that  $S'S'' = S$ , that would make  $C(j_o) = C(j^1)$ : contradiction.
- More generally, given two sub-streams  $S', S''$ , the sub-stream dominant element of  $S'S''$  is one of the sub-stream dominants in  $S'$  or  $S''$  and its frequencies are defined by

$$e - k - 1 \text{ in } S', \frac{n}{2} + 1 - (e - k - 1) \text{ in } S''$$

where  $e$  is an arbitrary number of appearances of  $j^2 : e < \frac{n}{2} + 1$  The reader will see as the  $+1$  in the right side guarantees the sub-stream dominant to be maximum. The sides can be swapped without hurting generality.

In order to better illustrate, we provide an algorithm (the Boyer–Moore majority vote algorithm) that exploits the sub-stream dominance. Our idea is to *go up* with our counter when we are fed an object we've already met, and to *go down* when we are fed an object different from us. Once we reach the bottom (0), we know by 7 that we either reached zero for an object  $\neq j^1$ , and we can ignore it, or we reached zero for  $j^1$ , that being the most frequent element will have sub-stream dominance in the remaining sequence, which means it will *go upper* than any other element in that sub-stream. Given that any other element has lower frequency, the most frequent of them,  $j^2$  will *go down*  $e < \frac{n}{2} + 1$  times.

```

1: function MAJORITY_VOTE_ALGORITHM( $S$ ):
2:    $candidate \leftarrow \text{null}$ 
3:    $counter \leftarrow 0$ 
4:   for all  $object \in S$  do
5:     if  $counter = 0$  then
6:        $candidate \leftarrow object$                                  $\triangleright$   $candidate$  is no more dominant, swap it
7:     else if  $object = candidate$  then
8:        $counter \leftarrow counter + 1$ 
9:     else if  $object \neq candidate$  then
10:       $counter \leftarrow counter - 1$ 
11:  return  $candidate$ 

```

## 8 Count-min sketch: extension to negative counters

Check the analysis seen in class, and discuss how to allow  $F[i]$  to change by arbitrary values read in the stream. Namely, the stream is a sequence of pairs of elements, where the first element indicates the item  $i$  whose counter is to be changed, and the second element is the amount  $v$  of that change ( $v$  can vary in each pair). In this way, the operation on the counter becomes  $F[i] = F[i] + v$ , where the increment and decrement can be now seen as  $(i, 1)$  and  $(i, -1)$ .

### 8.1 First solution

Let  $F$  be the implicit vector of size  $n$  that receives updates  $(i, v)$  such that  $F[i] = F[i] + v$  (initially,  $F[i] = 0 \forall i \in [1, n]$ ). The Count-Min (CM) sketch data structure, with parameters  $\varepsilon$  and  $\delta$ , consists of:

- a table  $T$  of size  $r \times c$ , where  $r = \log \frac{1}{\delta}$  and  $c = \frac{\varepsilon}{\varepsilon}$ , that we use to approximate queries on  $F$ .
- $r$  hash functions  $h_1, \dots, h_r$ , chosen uniformly at random from a pairwise-independent family, that maps from  $\{1, \dots, n\}$  to  $\{1, \dots, c\}$ .

The parameters specify that the error in answering a query is in within a factor of  $\varepsilon$  with probability  $1 - \delta$ . When an update  $(i, v)$  arrives, the table is updated as follows:

$$T[j, h_j(i)] = T[j, h_j(i)] + v \quad \forall j = 1, \dots, r$$

A point query returns an approximation  $\tilde{F}[i]$  of  $F[i]$ . If we assume that  $F[i] \geq 0$ , then:

$$\tilde{F}[i] = \min_{j \in [1, r]} T[j, h_j(i)]$$

with the guarantees:

1.  $\tilde{F}[i] \geq F[i]$
2.  $\mathbb{P} \left[ \tilde{F}[i] > F[i] + \varepsilon \|F\| \right] \leq \delta$ .

The problem statement asks to remove the assumption  $F[i] \geq 0$ . In this case the point query becomes:

$$\tilde{F}[i] = \text{median}_{j \in [1, r]} T[j, h_j(i)]$$

with the guarantee that  $\mathbb{P} \left[ F[i] - 3\varepsilon \|F\| \leq \tilde{F}[i] \leq F[i] + 3\varepsilon \|F\| \right] < 1 - \delta^{1/4}$ .

*Proof.* Let  $I_{i,j,k}$  the indicator variable that is 1 if  $i \neq k \wedge h_j(i) = h_j(k)$  and 0 otherwise. Observe that, by pairwise independence of the hash functions,

$$\mathbb{E} [I_{i,j,k}] = \mathbb{P} [h_j(i) = h_j(k)] \leq \frac{1}{c} = \frac{\varepsilon}{e}$$

and, given  $X_{i,j} = \sum_{k=1}^n I_{i,j,k} F[k]$

$$\mathbb{E} [|X_{i,j}|] = \mathbb{E} \left[ \sum_{k=1}^n I_{i,j,k} |F[k]| \right] \leq \sum_{k=1}^n |F[k]| \mathbb{E} [I_{i,j,k}] \leq \frac{\varepsilon}{e} \|F\| \quad (13)$$

The probability that any count (estimate of  $F[i]$ ) is off by more than  $3\varepsilon \|F\|$  is:

$$\begin{aligned} & \mathbb{P} [|T[j, h_j(i)] - F[i]| \geq 3\varepsilon \|F\|] \\ &= \mathbb{P} [|F[i] + X_{i,j} - F[i]| \geq 3\varepsilon \|F\|] && \text{by construction of } T \\ &= \mathbb{P} [|X_{i,j}| \geq 3\varepsilon \|F\|] \\ &\leq \frac{\mathbb{E} [|X_{i,j}|]}{3\varepsilon \|F\|} && \text{by the Markov inequality} \\ &\leq \frac{\varepsilon \|F\|}{e} \frac{1}{3\varepsilon \|F\|} && \text{by (13)} \\ &= \frac{1}{3e} < \frac{1}{8} \end{aligned}$$

The median is a bad choice if more than half of the  $r = \ln 1/\delta$  estimates are bad. We define  $Y$  to be the number of bad estimates:  $Y = \sum_{j=1}^r Y_j$ , where  $Y_j$  is 1 if  $|X_{i,j}| \geq 3\epsilon\|F\|$  and 0 otherwise. Hence the median is a bad choice if the sum  $Y$  is at least  $r/2$ , and this happens with probability:

$$\mathbb{P}\left[Y > \frac{1}{2} \ln \frac{1}{\delta}\right]$$

We can apply the Theorem 1 (with  $\mu = \mathbb{E}[Y] = \sum_{j=1}^r \mathbb{P}[j \text{ is a bad estimate}] < r/8$ ) to upper bound this probability.  $\square$

## 8.2 Second solution

We trivially have some changes over  $X_{ji}$  and  $\tilde{F}[i]$ :

$$\begin{aligned} X_{ji} &= \Sigma^n(I_k F[k]) \\ \tilde{F}[i] &= F[i] + X_{ji} \end{aligned}$$

$X_{ji}$  can vary as complementary increments  $(+v_i, +v_j, +v_k, -v_k, -v_j, -v_i)$  can make it  $\leq 0$  without necessarily being updates on  $i$ . In order to have a minimum error estimate we'll pick the *median* value of  $X_{ji}$ .

$\tilde{F}[i] = F[i] + X_{ji}$  by the above assertion the counter  $\tilde{F}[i]$  could have no garbage, or negative garbage according to the other increments. Therefore by choosing the minimum  $\tilde{F}[i]$  over the  $\log(\frac{1}{\delta})$  we could actually pick the most perturbed result (think of a collision with the biggest negative increment over one row).

The last step is the probability proof:  $\mathbb{P}[\tilde{F}[i] > F[i] + \epsilon\|F\|]$ . Since by 8.2  $X_{ji}$  can be either positive or negative we pick the *median* value  $med$  over the minimum  $m$  and maximum  $M$ . We are then able to split our error analysis in two of equivalent size  $\frac{r}{2}$ :

$$\begin{aligned} \mathbb{P}[\tilde{F}[i] \in F[i] + \epsilon\|F\|] &= \\ \mathbb{P}[\tilde{F}[i] \geq F[i] - |\epsilon\|F\| \wedge \tilde{F}[i] \leq F[i] + |\epsilon\|F\|] &= \\ \mathbb{P}[\tilde{F}[i] \geq F[i] - |\epsilon\|F\|) \mathbb{P}[\tilde{F}[i] \leq F[i] + |\epsilon\|F\|] \end{aligned}$$

Let us define  $p_0 = \mathbb{P}[\tilde{F}[i] \geq F[i] - |\epsilon\|F\|]$  and  $p_1 = \mathbb{P}[\tilde{F}[i] \leq F[i] + |\epsilon\|F\|]$ . We can report the analysis seen in class adjusted for the number of rows  $\frac{r}{2}$

$$\begin{aligned} p_0 &= \prod_{\frac{r}{2}} \mathbb{P}[|X_{ji}| \geq -|\epsilon\|F\|] = \\ &= -\frac{1}{2^{\frac{r}{2}}} = -\delta^{\frac{1}{2}} \end{aligned}$$

Now for  $p_1$ :

$$\begin{aligned} p_1 &= \mathbb{P}[\tilde{F}[i] \leq F[i] + |\epsilon\|F\|] = \\ 1 - \mathbb{P}[\tilde{F}[i] \geq F[i] + |\epsilon\|F\|] &= 1 - \frac{1}{2^{\frac{r}{2}}} = 1 - \delta^{\frac{1}{2}} \end{aligned}$$

Giving us a combined probability of  $p = p_0 p_1 = (-\delta^{\frac{1}{2}})(1 - \delta^{\frac{1}{2}}) = -\sqrt{\delta} + \delta$ .

## 9 Count-min sketch: range queries

Show and analyze the application of count-min sketch to range queries  $(i, j)$  for computing  $\sum_{k=i}^j F[k]$ . Hint: reduce the latter query to the estimate of just  $t \leq 2 \log n$  counters  $c_1, c_2, \dots, c_t$ . Note that in order to obtain a probability at most  $\delta$  of error (i.e. that  $\sum_{l=1}^t c_l > \sum_{k=i}^j F[k] + 2\epsilon \log n \|F\|$ ), it does not suffice to say that it is at most  $\delta$  the probability of error of each counter  $c_l$ : while each counter is still the actual wanted value plus the residual as before, it is better to consider the sum  $V$  of these  $t$  wanted values and the sum  $X$  of these residuals, and apply Markov's inequality to  $V$  and  $X$  rather than on the individual counters.

**Solution.** A range  $[a, b]$  is a dyadic range if its length is a power of two ( $l = 2^y$ ), and begins at a multiple of its own length:  $[j2^y + 1, (j+1)2^y]$ . For example,  $[13, 16]$  can be written as  $[3 \cdot 2^2 + 1, (3+1) \cdot 2^2]$ . Any arbitrary range of size  $s$  can be partitioned into  $O(\log s)$  dyadic ranges, for example [Cormode11]:

$$[18, 38] = [18, 18] \cup [19, 20] \cup [21, 24] \cup [25, 32] \cup [33, 36] \cup [37, 38]$$

Let  $n$  be the size of the implicit vector  $F$  whose entries we want to approximate. The idea is to maintain a collection  $C$  of  $\log_2 n$  CM sketches, one for each set of dyadic ranges of length  $2^y \forall y \in [0, \log_2 n - 1]$ . The operations on  $C$  becomes:

- **Update** ( $F[i] = F[i] + v$ ). Every sketch in  $C$  is updated, since each point  $1 \leq i \leq n$  is member of  $\log_2 n$  dyadic ranges.
- **Range queries** ( $\sum_{i=l}^r F[i]$ ). The range  $[l, r]$  is partitioned into at most  $2 \log_2 n$  dyadic ranges. For each partition, a point query is made to the corresponding sketch in  $C$ ; the (estimated) result of the range query is the sum of the point queries. See Figure 1.

The time to compute the estimate or to make an update is  $O(\log n \log \frac{1}{\delta})$ . The space used is  $O(\frac{1}{\epsilon} \log n \log \frac{1}{\delta})$ , because each sketch requires  $O(\frac{\epsilon}{\delta} \ln \frac{1}{\delta})$  space [Cormode05].

Let  $F[l..r] = \sum_{i=l}^r F[i]$  be the answer to the range query and  $\tilde{F}[l..r]$  the estimate. The guarantees are:

- $\tilde{F}[l..r] \geq F[l..r]$
- $\mathbb{P} \left[ \tilde{F}[l..r] > F[l..r] + 2\epsilon \log n \|F\| \right] \leq \delta$

*Proof.* Let  $I_{i,j,k}$  the indicator variable that is 1 if  $i \neq k \wedge h_j(i) = h_j(k)$  and 0 otherwise. From the analysis of CM sketch, we know that  $\mathbb{E}[I_{i,j,k}] \leq \frac{\epsilon}{e}$  and

$$\mathbb{E}[X_{i,j}] = \mathbb{E} \left[ \sum_{k=1}^n I_{i,j,k} F[k] \right] \leq \frac{\epsilon}{e} \|F\|$$

A range query  $\tilde{F}[l..r]$  is assembled by  $t \leq 2 \log n$  point queries on dyadic intervals  $[l_1, r_1], [l_2, r_2], \dots, [l_t, r_t]$ . The expectation of the additive error  $X^j$  for any estimator (i.e. the  $j$ -th row of any of the  $t$  CM sketches queried) is

$$\mathbb{E}[X^j] = \mathbb{E} \left[ \sum_{s=1}^t X_{i,j} \right] = \sum_{s=1}^t \mathbb{E}[X_{i,j}] \leq (2 \log n) \frac{\epsilon}{e} \|F\| \quad (14)$$

Also observe that the  $j$ -th estimation of  $F[l..r]$  is

$$\tilde{F}^j[l..r] = \sum_{s=1}^t \tilde{F}^j[l_s..r_s] = \sum_{s=1}^t (F[l_s..r_s] + X_{i,j}) = F[l..r] + X^j \quad (15)$$

Now, for the  $j$ -th estimation we can compute

$$\begin{aligned}
& \mathbb{P} \left[ \tilde{F}^j[l..r] > F[l..r] + 2\varepsilon \log n \|F\| \right] \\
&= \mathbb{P} \left[ X^j > 2\varepsilon \log n \|F\| \right] && \text{because of (15)} \\
&\leq \frac{\mathbb{E} [X^j]}{2\varepsilon \log n \|F\|} && \text{by the Markov inequality} \\
&\leq \frac{2\varepsilon \log n \|F\|}{e} \frac{1}{2\varepsilon \log n \|F\|} && \text{because of (14)} \\
&= e^{-1}
\end{aligned}$$

Since we have  $\ln \frac{1}{\delta}$  independent hash functions  $\prod_{j=1}^{\ln(1/\delta)} e^{-1} = e^{-\ln(1/\delta)} = \delta$  □

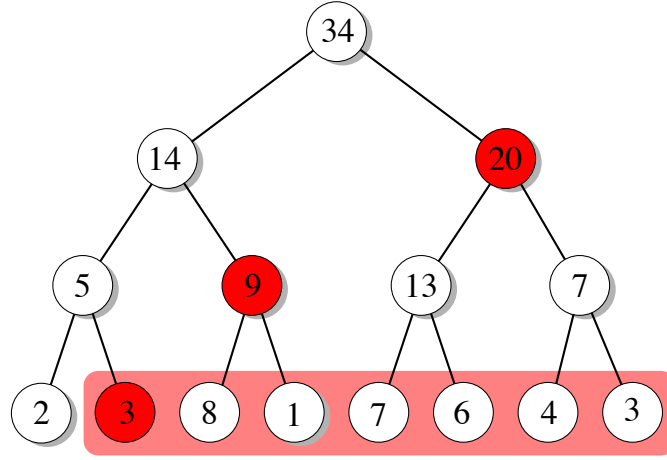


Figure 1: A hierarchy of dyadic ranges. The leaves are the ranges of length  $2^0$ , while the root corresponds to the single range of length  $2^3$ . Each level of the tree can be seen as a CM sketch table. To estimate  $\sum_{k=2}^8 F[k]$ , the range  $[2, 8]$  is decomposed into dyadic ranges  $[2, 2], [3, 4], [5, 8]$ . Each node contains the sum of the values stored in its children. Red nodes are queried and their sum is returned. Adapted from [Cormode11].

## 10 Space-efficient perfect hash

Consider the two-level perfect hash tables presented in [CLRS] and discussed in class. As already discussed, for a given set of  $n$  keys from the universe  $U$ , a random universal hash function  $h : U \rightarrow [m]$  is employed where  $m = n$ , thus creating  $n$  buckets of size  $n_j \geq 0$ , where  $\sum_{j=0}^{n-1} n_j = n$ . Each bucket  $j$  uses a random universal hash function  $h_j : U \rightarrow [m_j]$  with  $m_j = n_j^2$ . Key  $x$  is thus stored in position  $h_j(x)$  of the table for bucket  $j$ , where  $j = h(x)$ .

This problem asks to replace each such table by a bitvector of length  $n_j^2$ , initialized to all 0s, where key  $x$  is discarded and, in its place, a bit 1 is set in position  $h_j(x)$  (a similar thing was proposed in Problem 4 and thus we can have a one-side error). Design a space-efficient implementation of this variation of perfect hash, using a couple of tips. First, it can be convenient to represent the value of the table size in unary (i.e.,  $x$  zeroes followed by one for size  $x$ , so 000001 represents  $x = 5$  and 1 represents  $x = 0$ ). Second, it can be useful to employ a rank-select data structure that, given any bit vector  $B$  of  $b$  bits, uses additional  $o(b)$  bits to support in  $O(1)$  time the following operations on  $B$ :

- $rank_1(i)$ : return the number of 1s appearing in the first  $i$  bits of  $B$ .
- $select_1(j)$ : return the position  $i$  of the  $j$ th 1, if any, appearing in  $B$  (i.e.  $B[i] = 1$  and  $rank_1(i) = j$ ).

Operations  $rank_0(i)$  and  $select_0(j)$  can be defined in the same way as above. Also, note that  $o(b)$  stands for any asymptotic cost that is smaller than  $\Theta(b)$  for  $b \rightarrow \infty$ .

**Solution.** The two-level hashing layers are comprised of:

1. A hash function  $h$  and  $n$  pointers, each addressing one bitvector  $B_j$  of size  $n_j^2$ .
2.  $n$  hash functions  $h_j$  hashing to their relative bitvector  $B_j$  of size  $n_j^2$ .

Instead of storing a vector of pointers to the secondary hash tables, we merge the bitvectors  $B_j$ , one after the other, in a flat array  $B = B_0 B_1 \cdots B_{n-1}$ . Note that, except for the  $n$  pointers removed at the first level, the space is neither reduced nor increased, as no further bits are used to separate two adjacent buckets. We then store the size of each  $B_j$  in unary in an auxiliary bitvector  $L$ :

$$L = \underbrace{00 \dots 0 1}_{n_0^2 \text{ times}} \underbrace{00 \dots 0 1}_{n_1^2 \text{ times}} \cdots \underbrace{00 \dots 0 1}_{n_{n-1}^2 \text{ times}}$$

$L$  is associated with a rank-select data structure.

**Queries.** The starting index of  $B_j$  in  $B$  is computed in  $O(1)$  with the following function on  $L$ :

$$\phi(j) = rank_0(select_1(j))$$

This operation:

1. finds the position of the  $j$ th 1 with  $select_1(j)$ , that is, the index in  $L$  that precedes the start of the unary representation of  $n_j^2$ ;
2. calculates the sum of the sizes of the preceding bitvectors ( $\sum_{i=0}^{j-1} n_i^2$ ) by computing the number of 0s with  $rank_0(select_1(j))$  (this is the starting position of the desired bitvector).

We can determine (with a probability of error) whether a key  $k$  belongs to the set  $S \subset U$ , by testing whether  $B[i + h_{h(k)}(k)]$  is equal to 1, where  $i = \phi(h(k))$  is the starting position of the  $h(k)$ th bitvector, and  $h_{h(k)}(k)$  is the offset for the secondary level.

**Hash functions space optimization.** We now try to improve the space for the hash functions parameters  $a_j, b_j, p_j$ . First, we select the lowest  $p_j$  possible, that is, the first  $p_j > n_j^2$ : by Bertrand postulate such a prime  $p_j$  exists for  $n_j^2 \leq p_j \leq 2n_j^2 - 2$ . We therefore need a space of at most  $2n_j^2$  bits for each parameter in an unary encoding. Since we have  $a \in \mathbb{Z}_{p_j}^*, b \in \mathbb{Z}_{p_j}$ , we have a space of  $3 \cdot 2n_j^2$  for a triple  $(a_j, b_j, p_j)$  of any given hash function. Let us now compute the space for the  $n$  hash functions:

$$Y = 3n + 3 \sum_{j=0}^{n-1} 2n_j^2 = 3n + 6 \sum_{j=0}^{n-1} n_j^2$$

where  $3n$  is the number of 1s that separate the unary encoding of the parameters. We have also an auxiliary space  $o(Y)$  due to the rank-select data structure.

**Space.** The space occupied by the whole data structure is:

- $X = \sum_{j=0}^{n-1} n_j^2$  bits for  $B$ ;
- $X + n$  bits for  $L$ , since  $X$  is the number of 0s and  $n$  is the number of 1s in  $L$ ;
- $o(X + n)$  bits for the rank-select auxiliary data structure for  $L$ ;
- $Y + o(Y)$  bits for the hash functions, as shown in the previous paragraph.

Since  $\mathbb{E} \left[ \sum_{j=0}^{n-1} n_j^2 \right] < 2n$ , as shown in the perfect hashing analysis [Cormen09], the total space is  $O(n)$ .



## 11 Bloom filters vs. space-efficient perfect hash

Recall that classic Bloom filters use roughly  $1.44 \log_2(1/f)$  bits per key, as seen in class (where  $f = (1-p)^k$  is the failure probability minimized for  $p \approx e^{-\frac{kn}{m}} = 1/2$ ). The problem asks to extend the implementation required in Problem 10 by employing an additional random universal hash function  $s : U \rightarrow [m]$  with  $m = \lceil 1/f \rceil$ , called signature, so that  $s(x)$  is also stored (in place of  $x$ , which is discarded). The resulting space-efficient perfect hash table  $T$  has now a one-side error with failure probability of roughly  $f$ , as in Bloom filters: say why. Design a space-efficient efficient implementation of  $T$ , and compare the number of bits per key required by  $T$  with that required by Bloom filters.

**Solution.** We extend the data structure discussed in Section 10 with an additional array  $C$  that stores the  $n$  signatures. Leveraging the fact that  $B$  has as many 1s as the number of signatures we need to store, we fill the array  $C$  with the signatures in this way: the entry  $C[i]$  contains the signature  $s(k)$  in binary if and only if  $k$  is the key that caused a the  $i$ th one in  $B$ . We also need an additional space of  $o(X)$  to perform rank-select operations on  $B$  in  $O(1)$ .

With this new data structure, given a key  $l \in U$ , to check whether it belongs to  $S$  we check whether  $B[i] = 1$  where  $i = \phi(h(l)) + h_{h(l)}(l)$  as before, but also:

$$C[\text{rank}_1(i) - 1] \stackrel{?}{=} s(l). \quad (16)$$

In fact,  $\text{rank}_1(i) - 1$ , executed on  $B$ , returns the position in  $C$  where we find the signature for the key  $k \in S$  that caused  $B[i] = 1$ .

We have an error whenever  $l \notin S$ , but (16) is true. Since the signature has length  $1/f$ ,  $\Pr(\text{error}) \leq \frac{1}{1/f} = f$ .

**Bits per key comparison.** If we consider only the space used by  $C$ ,  $B$  and  $L$ , our solutions uses  $n \log_2 \frac{1}{f} + X + (X + n)$  bits to store  $n$  keys. Since  $\mathbb{E}[X] < 2n$ , on average we need approximately  $\log_2 \frac{1}{f} + 5$  bits per key, that compared to the Bloom filters is better for  $f < 2^{-\frac{5}{0.44}}$ .

## 12 MinHash sketches

As discussed in class, for a min-wise independent family  $\mathcal{H}$ , we can associate a sketch

$$s(X) = \langle \min h_1(X), \min h_2(X), \dots, \min h_k(X) \rangle$$

with each set  $X$  in the given data collection, where  $h_1, h_2, \dots, h_k$  are independently chosen at random from  $\mathcal{H}$ . Consider now any two sets  $A$  and  $B$ , with their sketches  $s(A)$  and  $s(B)$ . Can you compute a sketch for  $A \cup B$  using just  $s(A)$  and  $s(B)$  in  $O(k)$  time? Can you prove that it is equivalent to compute  $s(A \cup B)$  from scratch directly from  $A \cup B$ ?

**Solution.** We claim that

$$\langle \min(\min h_1(A), \min h_1(B)), \dots, \min(\min h_k(A), \min h_k(B)) \rangle$$

which can be computed from  $s(A)$  and  $s(B)$  with  $\Theta(k)$  comparisons, is equivalent to

$$s(A \cup B) = \langle \min h_1(A \cup B), \dots, \min h_k(A \cup B) \rangle$$

In fact, note that  $\forall i \in [1, k]$ ,

$$\min h_i(A \cup B) = \min(h_i(A) \cup h_i(B)) = \min(\min h_i(A), \min h_i(B))$$

The second equality is a trivial property of the union, for  $h_i(A \cup B) = h_i(A) \cup h_i(B)$  we give the following

*Proof.* First we prove that  $h_i(A \cup B) \subseteq h_i(A) \cup h_i(B)$ :

$$l \in h_i(A \cup B) \implies \exists s \in A \cup B \text{ such that } h_i(s) = l$$

Since  $s \in A \cup B \implies s \in A \vee s \in B$ , we have two cases: if  $s \in A$ , then  $h_i(s) \in h_i(A)$ , hence  $h_i(s) \in h_i(A) \cup h_i(B)$ ; if instead  $s \in B$ , then  $h_i(s) \in h_i(B)$ , hence  $h_i(s) \in h_i(A) \cup h_i(B)$ .

Now we prove  $h_i(A) \cup h_i(B) \subseteq h_i(A \cup B)$ :

$$l \in h_i(A) \cup h_i(B) \implies l \in h_i(A) \vee l \in h_i(B)$$

If  $l \in h_i(A)$ , then  $\exists s \in A$  such that  $h_i(s) = l$ ; since  $s \in A \implies s \in A \cup B$ , we have  $h_i(s) \in h_i(A \cup B)$ . If instead  $l \in h_i(B)$ , then  $\exists s \in B$  such that  $h_i(s) = l$ ; since  $s \in B \implies s \in A \cup B$ , we have  $h_i(s) \in h_i(A \cup B)$ .  $\square$

## 13 Randomized min-cut algorithm

Consider the randomized min-cut algorithm discussed in class. We have seen that its probability of success is at least  $1/\binom{n}{2}$ , where  $n$  is the number of its vertices.

- Describe how to implement the algorithm when the graph is represented by adjacency lists, and analyze its running time. In particular, a contraction step can be done in  $O(n)$  time.
- A weighted graph has a weight  $w(e)$  on each edge  $e$ , which is a positive real number. The min-cut in this case is meant to be min-weighted cut, where the sum of the weights in the cut edges is minimum. Describe how to extend the algorithm to weighted graphs, and show that the probability of success is still  $\geq 1/\binom{n}{2}$  [hint: define the weighted degree of a node].
- Show that running the algorithm multiple times independently at random, and taking the minimum among the min-cuts thus produced, the probability of success can be made at least  $1 - 1/n^c$  for a constant  $c > 0$  (hence, with high probability).

**Contraction step.** We assume that the adjacency lists  $Adj[x]$  are sorted  $\forall x \in V$ . We also note that the multigraph is undirected, therefore  $v \in Adj[u]$  if and only if  $u \in Adj[v]$ . We maintain an attribute  $pe$  in each edge to store the number of parallel edges.

To contract an edge  $(u, v)$ , we have to merge the two adjacency lists  $Adj[u]$  and  $Adj[v]$  into a single sorted adjacency list  $Adj[uv]$  — like the merge procedure in merge sort — ensuring that:

- if the current node in  $Adj[u]$  is  $v$ , skip the node, since it will not appear in  $Adj[uv]$  (do the same with  $Adj[v]$  and  $u$ );
- if the current nodes are equal to  $x$ , add  $x$  to  $Adj[uv]$  and set  $(uv, x).pe = (u, x).pe + (v, x).pe$ .

Finally, we replace  $Adj[u]$  and  $Adj[v]$  with the newly created  $Adj[uv]$ . The total cost of this step is  $O(n)$ .

**Weighted graph extension.** We now extend the above to weighted multi-graph by re-conducting it to the non-weighted case. First we'll define the *min-weighted cut*, the weighted cut where  $\min_{|c|>0} C = \min_C \sum \omega(c)$ , that is the sum of the weights determines the weight of the overall cut. We then have the weight for the overall graph:

$$\omega_G = \sum^n (\omega(n))$$

and for the min cut:  $\omega_{min}$ . Similarly as we've seen in class we can define  $\omega_G$  in function of  $\omega_{min}$ :

$$\omega_G = \frac{\sum^n (\omega(n))}{2} = \frac{n \cdot \omega_{min}}{2}$$

Giving us an error probability of

$$\Pr(error) = \frac{\text{WEIGHT}(\text{MIN-CUT})}{\text{WEIGHT}(\text{GRAPH})} = \frac{\omega_{min}}{\omega_G} = \frac{\omega_{min}}{\frac{n \cdot \omega_{min}}{2}} = \frac{2}{n}$$

That is, we reach the same error probability seen in class. Therefore we can safely assume that the analysis is the same.

The algorithm does not need to be edited but in the computation of the min cut, which is now  $\omega_c$  and not  $\sum_{e^{i,j}} (e_i)$  where  $e_i$  is an edge between  $i, j$  last remaining nodes in the cut.

**Error probability.** By the analysis seen in class we have an error probability of

$$1 - \Pr(\text{success}) = 1 - \frac{1}{\binom{n}{1}}$$

if we then run the algorithm some  $d \cdot \frac{1}{\binom{n}{2}}$  times, the probability of success becomes

$$1 - \left(1 - \frac{1}{\binom{n}{2}}\right)^{c \cdot \frac{1}{\binom{n}{2}}} \geq 1 - e^{-d}$$

by  $d = c \ln(c)$  we have an error probability of  $\leq \frac{1}{n^c}$ .

## 14 External memory implicit searching

Given a static input array  $A$  of  $N$  keys in the EMM (external memory or cache-aware model), describe how to organize the keys inside  $A$  by suitably permuting them during a preprocessing step, so that any subsequent search of a key requires  $O(\log_B N)$  block transfers using just  $O(1)$  memory words of auxiliary storage (besides those necessary to store  $A$ ). Clearly, the CPU complexity should remain  $O(\log N)$ . Discuss the I/O complexity of the above preprocessing, assuming that it can use  $O(N/B)$  blocks of auxiliary storage. (Note that the additional  $O(N/B)$  blocks are employed only during the preprocessing; after that, they are discarded as the search is implicit and thus just  $O(1)$  words can be employed.)

**Solution.** The idea is to construct a B-tree, a balanced search tree where each node has  $B$  keys. The keys inside a node  $x$  divide the interval of keys stored below  $x$  in  $B + 1$  intervals, therefore each node has  $B + 1$  children. We don't store pointers explicitly, the index of the  $j$ -th child of a node  $i$  is

$$i(B + 1) + B(j + 1) \quad \text{where } 1 \leq j \leq B + 1 \text{ and } 0 \leq i < A.length$$

This is a generalization of the formula for implicit binary heaps, in which  $B = 1$ .

Assuming that  $A$  is sorted, we construct the tree as follows:

1. if  $A.length \leq B$  then STOP, since  $A$  is the root of the tree;
2. otherwise, select the keys in  $A$  to move to the upper level, those whose position  $i$  is such that  $i \bmod (B + 1) = B$ ;
3. let  $L$  be the keys selected in the previous step, store the remaining  $A \setminus L$  keys as leaves, sorted and grouped in blocks of size  $B$ ;
4. repeat the process with  $A \leftarrow L$ .

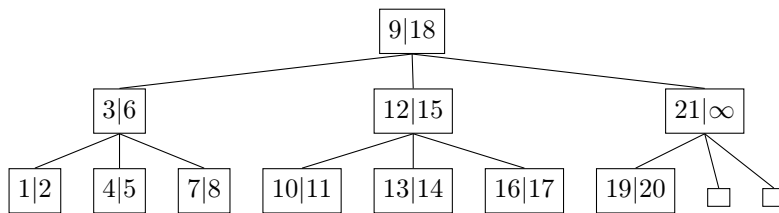
For example, suppose that  $B = 2$  and the keys are  $1, 2, \dots, 21$ . First, we select the keys to move to the upper level (those boxed), while the remaining keys will be the leaves of the tree:

1 2 3 4 5 6 7 8 9 10 11 12 13 14 15 16 17 18 19 20 21

We repeat the process with the keys selected at the previous iteration (again, the remaining keys form a new level of the tree):

3 6 9 12 15 18 21

In the third iteration we have 2 keys and we can stop, since they can be both stored in a single block that will be the root of the tree:



The tree will be represented in a file in BFS layout:

9 18 3 6 12 15 21 ∞ 1 2 4 5 7 8 10 11 13 14 16 17 19 20

**Space and I/O complexity.** The construction algorithm copies the permutation of the  $N$  keys from  $A$  to a new file, and thus the total additional space needed in external memory is  $O(N/B)$  blocks. After the preprocessing,  $A$  is discarded, hence the auxiliary space is  $O(1)$ .

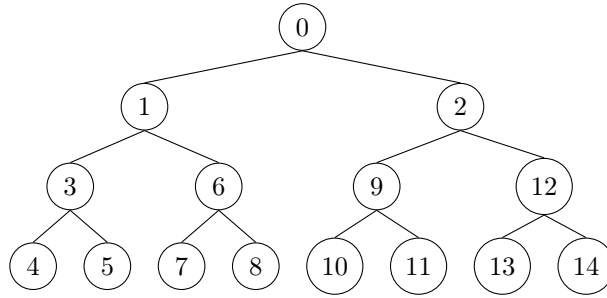
The I/O complexity of the step 3 in the first iteration is  $O(N/B)$ , because we move  $m = N - N/(B+1)$  keys in  $A \setminus L$ , from  $A$  to the new file, with  $O(m/B) = O(N/B)$  read/write operations. Step 1 requires only  $O(1)$  transfers to write the root block to the beginning of the new file. Steps 2 and 4 don't require any transfer from/to external memory, only CPU operations. Subsequent iterations of the algorithm work on smaller portions of  $A$ , thus the first iteration dominates with a total cost of  $O(N/B)$  I/O operations.

## 15 Implicit navigation in vEB layout

Consider  $N = 2^h - 1$  keys where  $h$  is a power of 2, and the implicit cache-oblivious vEB layout of their corresponding complete binary tree, where the keys are suitably permuted and stored in an array of length  $N$  without using pointers (as it happens in the classical implicit binary heap but the rule here is different). The root is in the first position of the array. Find a rule that, given the position of the current node, it is possible to locate in the array the positions of its left and right children. Discuss how to apply this layout to obtain (a) a static binary search tree and (b) a heap data structure, discussing the cache complexity.

### 15.1 Solution.

We assume that the keys are stored in a zero-indexed array  $T$ . Given an index  $0 \leq i < T.length$  we will give a rule for computing the index of its children.



As by hypothesis, we'll assume  $H = 2^k, k \in \mathbb{Z} \setminus \{0, 1\}$ .

**Shifted layouts** We start by introducing the notion of *shifted layout*: a *shifted layout*  $L'$  of a layout  $L$  is a layout of the same size whose root is shifted from the one of  $L$  of an offset  $g$ :

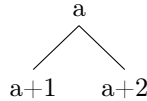
$$\text{HEIGHT}(L) == \text{HEIGHT}(L'), L'.\text{root} = L.\text{root} + a \quad (17)$$

Then

$$n \in L, n' \in L' : n.\text{index} == n'.\text{index} \rightarrow n'.\text{position} == n.\text{position} + a \quad \forall n \in L, n' \in L' \quad (18)$$

Trivially, as the two layouts are shifted in the root, the same shift is propagated to the remaining nodes in the layout.

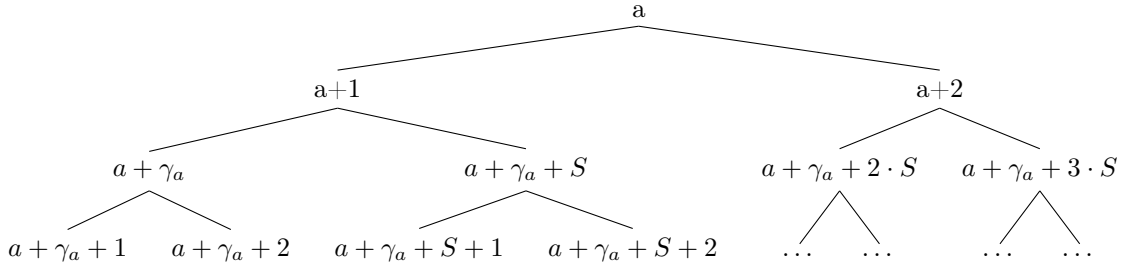
**Contiguous children** Given a layout  $L$ , we are able to *cut* it until we reach a base case of layout  $L'$  of height 2 and root  $a$ . It is trivial to see as this is populated as the following shows:



by the definition of base case for the *vEB construction*. Then, provided we know to be in the  $a$  index we are able to answer trivially with

$$\begin{aligned} \text{LEFT}(A) &= a + 1 \\ \text{RIGHT}(A) &= a + 2 \end{aligned}$$

**Non-contiguous children** Given a layout  $L$ , we are able to *cut* it until we reach a base case of



The first two levels are trivial by 15.1. From the second to the third we find instead a jump  $\gamma_a$ . By construction of the *layout* the position  $a + \gamma_a$  is populated right after  $a + 2$ : in this trivial case we have a jump of size 2. This holds also for  $a + \gamma_a + S$ , as it is populated right after  $a + \gamma_a + 2$  and so on and so forth on the same level. Given that  $\text{height}(T) = 4$  we have jumps  $\gamma_a = 2$ : this does not hold for any height: let  $H = 2h$  be the height of a *vEB tree* s.t.  $h > 1$ ,  $a$  be a node at level  $h + 1$ ,  $b$  be a node at level  $l \neq h + 1$ , then  $\gamma_a = 2^h - 1$  and  $\gamma_a > \gamma_b \forall b$ . By construction the jump  $\gamma$  and the gap  $S$  grow exponentially w.r.t. the level of the tree  $T$ : given that the *vEB* construction halves the height of the current tree at every recursion, such height is maximum at the very first recursion. Moreover, given that the sub-tree rooted in  $a + \gamma_a + S$  is populated contiguously,  $T_{a+\gamma_a} \in [a + \gamma_a, a + \gamma_a + 1, \dots, a + \gamma_a + S - 1]$ : that is given any  $n \in [a + \gamma_a, a + \gamma_a + S - 1]$  we can know in which subtree  $T_i$  rooted in  $a + \gamma_a + k \cdot S$   $n$  is contained:

$$t = n - (a + \gamma_a) / S \quad (19)$$

**Powers of two** Let  $l \in [1, H]$  be a level for the *vEB tree*  $T$  s.t. the first root of that cut  $a = 2^{2^k}$ . Then  $\exists b : b = 2^{2^{2k}}$  and  $b$  is the first root at level  $2l$ .

**Proof** Let  $H = 2h$  the whole *vEB tree* and  $T_0$  the first top cut from 1 to  $h$  and  $T_1$  the bottom cut from  $h + 1$  to  $2h$ . By construction  $T_0$  is populated contiguously with  $[1, 2^h]$ . The  $2^h$  leafs at  $h$  create  $2^{h+1}$  nodes, roots for  $T_1$ : given that the cut was of size  $h$  each of the  $2^{h+1}$  sub-trees has  $2^h - 1$  nodes, for a total of  $(2^{h+1}) \cdot (2^h - 1) = 2^{2h} = 2^H$ .

**Algorithm** We can now combine the above to create an algorithm to find the children of a node  $i$ :

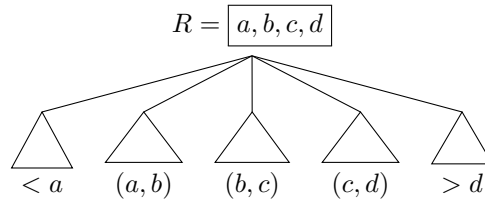
1. Find  $K = 2^{2^k}$  the largest power of a power of two lower than  $i$ : by 15.1 it will be in either the top or bottom half of  $T$ . Let  $l$  be the level of  $K$  and  $T'$  the respective cut.
2. By 15.1 find the sub-tree  $T'$  and relative root  $a$  with 19. Compute its root  $a = K + t \cdot (K - 1)$  by 19. By 15.1  $T'$  is a shifted tree of height  $h$  with offset  $K - 1$ : scale it back together with  $i$  and recursively apply on  $T'$ , cumulating the offsets and the levels as you go.
3. Once a base case of  $\text{height}(T') = 2$  we have two cases: i)  $i$  is the root, and we can return  $i + 1, i + 2$  by 15.1; ii)  $i$  is not the root and a jump  $\gamma_i$  has to be computed. By 15.1 we are able to do so, provided we have the level. Note as at step 2 of the algorithm we have also cumulated the levels, thus allowing us to do so.

## 16 1-D range query

Describe how to efficiently perform one-dimensional range queries for the data structures described in Problems 14 and 15. Given two keys  $k_1 \leq k_2$ , a range query asks to report all the keys  $k$  such that  $k_1 \leq k \leq k_2$ . Give an analysis of the cost of the proposed algorithm, asking yourself whether it is output-sensitive, namely, it takes  $O(\log_B N + R/B)$  block transfers where  $R$  is the number of reported keys.

### 16.1 First solution

**B-tree.** To report the keys in the given range we perform an inorder traversal of the tree starting from the root  $R$ : for each key  $k$  stored in the node  $R$  in position  $i$ , if  $k \geq k_1$ , then we traverse the  $i$ -th subtree of  $R$  and output  $k$ ; if  $k > k_2$ , then we traverse the  $i$ -th subtree of  $R$  and stop the for loop.



First, this algorithm searches for the key  $k_1$  and, in the worst case, it visits all the nodes in a walk from the root to a leaf with  $O(\log_B N)$  block transfers (see Figure 2). Afterward, it starts to report all the keys in the given range: this can be done with  $O(R/B)$  block transfers, provided that we can hold in memory one block for each level of the tree, which is stored contiguously in memory and scanned left-to-right. Consequently, the I/O complexity of the range query is  $O(\log_B N + R/B)$ .

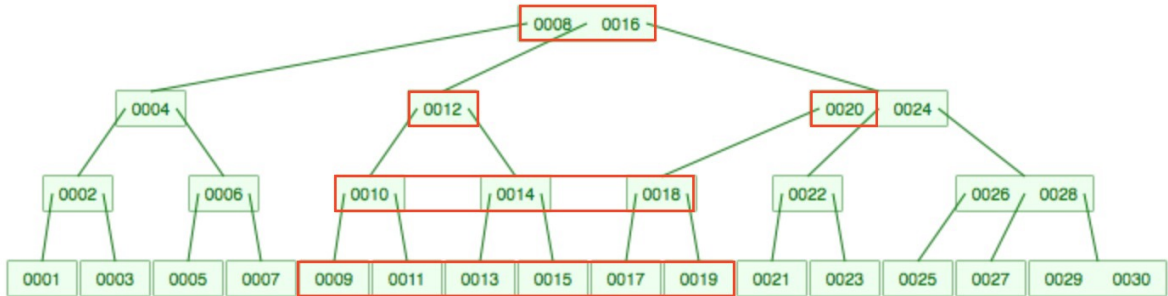


Figure 2: A B-tree with the keys in range  $[8, 20]$  highlighted in red.

### 16.2 Second solution

**B-tree.** By our construction a punctual query is answered with  $\log_B(N)$  memory transfers. Let us now extend such queries to ranges similarly to the previous exercise on range updates. Given  $k_1, k_2$  interval extremes and  $B_i$  the current block we store in memory  $S$ , sum of all the keys up until now, and  $I, J$  addresses of the block comprising the interval at some level  $k$  s.t.  $k$  is the highest level containing  $k_1, k_2$ . We assume we are provided with enough space to store also pointers to some, if not all the blocks that comprise  $k_1, k_2$  in order not to traverse them more than one time. If such storage is not available, we scale to pointers to the nodes comprising  $k_1, k_2$  and we restart our search from there if necessary. Given that our tree is a  $k$ -ary tree we search through a trivial  $k$ -ary search, summing the values we meet in the path if they are in the desired interval. Moreover our construction allows us to assert that given any sub-tree  $T'$  an in-order visit grants us a contiguous interval: we are then able to load each of the  $R'/B$  blocks in memory and sum them in  $S$ . If the interval is completed, then we have our answer in  $h' + \frac{R'}{B}$  loads; otherwise we need to go back to the previous highest node in  $p'$  where we split the interval and



repeat at most  $h'' + \frac{R''}{B}$ . Since we loaded only blocks whose keys belong to  $[k_1, k_2]$  we have  $R' + R'' = R$  and  $h'' + \frac{R''}{B} + \log_B h' + \frac{R'}{B} = 2 \log_B N + \frac{R}{B}$  giving an output-sensitive search algorithm.

**vEB layout.** We've shown previously as *vEB trees* with an height of  $2^k$  store adjacent indexes adjacently, effectively providing an ordered array that we can load with  $\frac{R}{B}$  blocks.

## 17 External memory mergesort

In the external-memory model (hereafter EM model), show how to implement the  $k$ -way merge (where  $(k+1)B \leq M$ ), namely, how to simultaneously merge  $k$  sorted sequences of total length  $N$ , with an I/O cost of  $O(\frac{N}{B})$  where  $B$  is the block transfer size. Also, try to minimize and analyze the CPU time cost.

**Solution.** Let

- $k = \frac{1}{4} \cdot \frac{M}{B}$  be the number of blocks the cache can hold.
- $B_s$  the blocks reserved in cache to store temporary merges.
- $N$  be the size of the problem.
- $B$  the block size, with  $B_i$  a pointer to the  $i^{th}$  block.
- $B_i^c$  the  $i^{th}$  block stored in memory.
- $\frac{N}{B}$  the number of blocks storing  $A$  in main memory.

We assume the sorted blocks we receive in input are ordered and use  $\frac{1}{3} \cdot \frac{M}{B}$  to store loaded blocks in a Fibonacci min-heap fashion.  $k$ -way mergesort operates on  $k$  sorted blocks we load in cache. We implement the merge operation as follows, outputting to external memory on file  $O$ :

```

1: function MERGE( $a, b$ ):
2:   for  $i \leftarrow 0; i < k; i++$  do
3:     load( $B_i, B_i^m$ )                                ▷ Load blocks in cache:  $O(\frac{N}{B})$ 
4:     busy_blocks =  $k - 1$                             ▷ Count busy blocks, that is blocks whose pointers
                                                         have not reached the end of the block
5:      $S \leftarrow \&S$                                   ▷ At most  $O(k \cdot B)$ , when every sorted block is greater
                                                         than the following
6:     while busy_blocks > 0 do
7:        $m \leftarrow \min\{\min B_i^c\}$                   ▷ Get the current minimum of every block in cache:  $O(k)$ 
8:        $B_j^c \leftarrow B_j^c + 1$                         ▷ Shift to next element in block  $B_j$  with current minimum element
9:       if  $B_j == EOB$  then
10:        busy_blocks  $\leftarrow$  busy_blocks - 1
11:       if busy_blocks == 1 then
12:         $S \leftarrow S :: flush(B_j^c)$                 ▷ Flush remaining block
13:         $*S \leftarrow m$ 
14:         $S \leftarrow S + 1$                             ▷ Increment storing blocks pointer
15:       write( $S, O$ )                                    ▷ Output

```

As we assumed, our blocks are ordered Fibonacci min-heaps, thus allowing us to merge them in  $k \cdot \log(n)$  time. We can then store  $S$  by finding the minimum element  $n$  times in constant time:  $n \cdot O(1) = O(n)$ , giving us a total cost of

$$k \cdot B + \log(k) \cdot O(k) + O(n)$$

respectively for construction, find the minimum over  $k$  Fibonacci heaps in the *merge* procedure and store it in memory.

## 18 External memory (EM) permuting

Given two input arrays  $A$  and  $\pi^{-1}$ , where  $A$  contains  $N$  elements and  $\pi^{-1}$  contains a permutation of  $\{1, \dots, N\}$ , describe and analyse an optimal external-memory algorithm for producing an output array  $C$  of  $N$  elements such that  $C[i] = A[\pi[i]]$  for  $1 \leq i \leq N$ .

**Solution.** We define an array  $\pi^{-1}$  s.t.  $\pi^{-1}[i] = j : \pi[j] = i$ , that is the inverse mapping of  $\pi$ : *position*  $\rightarrow$  *index*. We define an array  $A\pi^{-1}$  s.t.  $A\pi^{-1}[i] = (A[i], \pi^{-1}[i])$ : note as we need to load  $2 \cdot \frac{N}{B}$  blocks. We then sort  $A\pi^{-1}$  with the previously defined k-way mergesort in  $O(n \cdot \log(n))$  time and  $O(\frac{N}{B})$  loads. By filtering  $A\pi^{-1}$  over the first pair's element we have the ordered  $C$  we wanted to obtain, and we are able to write it to memory with  $O(\frac{N}{B})$  memory transfers. We then have a total of

$$m = 2 \cdot \frac{N}{B} + O(\frac{N}{B}) + O(\frac{N}{B}) \in O(\frac{N}{B})$$

$m$  memory transfers with complexity of  $O(N \cdot \log_k(N))$ .

## 19 Wrong greedy for minimum vertex cover

Find an example of (family of) graphs for which the following greedy approach fails to give a 2-approximation for the minimum vertex cover problem (and prove why this is so). Start out with an empty  $\tilde{S}$ . Choose each time a vertex  $v$  with the largest number of incident edges in the current graph. Add  $v$  to  $\tilde{S}$  and remove its incident edges. Repeat the process on the resulting graph as long as there are edges in it. Return  $|\tilde{S}|$  as the approximation of the minimal size of a vertex cover for the original input graph. Generalize your argument to show that the above greedy algorithm cannot actually provide an  $r$ -approximation for any given constant  $r > 1$ .

**Solution** We split the  $N = 3n$  nodes in two partitions of sizes  $n, 2n$  that we call  $L, R$  respectively. We then build the edges in the following way in order to obtain a minimum cover  $L$ :

1. If  $|R| = 0$  we stop.
2. Split the current slice  $R$  in half: let  $R^i$  be this slice.
3. For every node  $n \in R^i$  build  $i$  outgoing edges to each of the  $|L|$  nodes  $\in L$ . Note as for  $|R^0| = \frac{2n}{2} = n$  each node reaches exactly one node  $\in L$ . Halving  $R^0$  we double the number of outgoing edges to  $L$ , we maintain the number of outgoing edges from  $R$  constant. We can then repeat step 1 with  $R \leftarrow R \setminus R^i$ .

Once the above procedure ends, we have  $\deg(n) = \deg(n')$  for any  $n, n' \in L$  because at each split every node in  $L$  received an incoming edge. Each node in  $L$  will have an incoming edge for every node  $n \in R$ . The greedy algorithm then starts by removing the highest-degree node: we find this on top of  $R$ , as it was doubled at each iteration. It is therefore forced to reach  $R^0 = 1$  where  $\deg(n) = 1$ . As this was the very first cut, we have to remove  $n$  nodes. The set  $\tilde{S}$  has then  $\sum_{i=2}^n \frac{n}{i}$  nodes:

$$\sum_{i=2}^n \frac{n}{i} = n \sum_{i=2}^n \frac{1}{i} \approx n \ln n$$

Since  $n \ln n > n = \text{minimum cover}$ , we have a generalization for any  $n$ .

## A Hogwarts

The Hogwarts School<sup>1</sup> is modeled as a graph  $G = (V, E)$  where  $V$  is the set of castle's rooms and  $E \subseteq V \times V$  is the set of the stairs. Each stair is labelled with the time of appearance and disappearance, and can be walked in both directions, therefore the graph is undirected. The goal is to find, if possible, the minimum amount of time required to go from the first to the last room.

### A.1 Solution 1: Preprocessing-then-Dijkstra

Dijkstra is able to find the shortest path in a graph with non-negative weights on its edges. Our main idea is to create a Dijkstra compatible graph through a `NORMALIZE` function, then apply Dijkstra to it in order to find the shortest path. The core of the preprocessing is the `NORMALIZE` function which computes traversal times between nodes at a given time *time*:

```

1: function NORMALIZE(from, to, time):
2:    $t \leftarrow \infty$ 
3:   if  $start[v'] \leq t < end[v']$  then                                     ▷ No waiting time
4:      $t \leftarrow t + 1$ 
5:   else if  $t < start[v']$  then                                           ▷ Waiting time
6:      $t \leftarrow start[v'] + 1$ 
7:   else
8:      $t \leftarrow \infty$                                                  ▷ Available time already expired
9:   return  $t$ 

```

The normalize function is then applied to a node traversal:

#### A.1.1 Pseudo-code

```

1: create vertex set  $Q$  of unvisited nodes
2: create vertexes set  $E'$  of edges weight
3:  $time \leftarrow 0$                                                          ▷ Initial time for traversal
4:  $edges \leftarrow STAIRS\_OF(0)$                                          ▷ Get incoming/outgoing edges of the source node
5: function PROCESS(node, time)
6:   if  $edge \in visited\_edges$  then
7:     return
8:    $traversal\_time \leftarrow \infty$ 
9:   for all  $neighbor \in neighbors\_of\_node$  do
10:     $traversal\_time \leftarrow TRAVERSAL\_TIME(node, neighbor, time)$ 
11:     $E'[0][node] \leftarrow traversal\_time$                                 ▷  $E'[i][j]$  holds the weight/traversal
12:                                                         ▷ time for the stair between  $i$  and  $j$ 
13:    for all  $new\_neighbor \in neighbors\_of\_neighbor$  do
14:      NORMALIZE( $neighbor$ ,  $new\_neighbor$ ,  $traversal\_time$ )
15:   if DIJKSTRA( $V, E'$ ) =  $\infty$  then
16:     return -1
17:   else
18:     return  $t$ 

```

**Computational cost:**  $\Theta(n^2)$  if the vertex set in DIJKSTRA is implemented as an array.  $O(|E| + |V| \log |V|)$  with Fibonacci heap.

### A.2 Solution 2: HogwartsDijkstra

```

1: function HOGWARTSDIJKSTRA( $G$ ):
2:   create vertex set  $Q$  of unvisited nodes

```

<sup>1</sup>[http://didawiki.cli.di.unipi.it/lib/exe/fetch.php/magistraleinformatica/alg2/algo2\\_16/hogwarts.pdf](http://didawiki.cli.di.unipi.it/lib/exe/fetch.php/magistraleinformatica/alg2/algo2_16/hogwarts.pdf)

```

3:  for all vertex  $v \in V$  do ▷ initialization
4:       $time[v] \leftarrow \infty$  ▷ unknown time from source to  $v$ 
5:      add  $v$  to  $Q$  ▷ all nodes initially in  $Q$ 
6:   $time[0] \leftarrow 0$  ▷ time from source to source
7:  while  $Q \neq \emptyset$  do
8:       $u \leftarrow x \in Q$  with  $\min\{time[x]\}$ 
9:      remove  $u$  from  $Q$ 
10:     for all neighbor  $v$  of  $u$  do:
11:         if  $time[u] \leq appear[v]$  then
12:              $alt \leftarrow appear[v] + 1$  ▷ wait the appearance of the stair
13:         else if  $time[u] < disappear[v]$  then
14:              $alt \leftarrow time[u] + 1$  ▷ use the stair
15:         else
16:              $alt \leftarrow \infty$  ▷ the stair has already disappeared
17:         if  $alt < time[v]$  then
18:              $time[v] \leftarrow alt$  ▷ a quicker path to  $v$  has been found
19:  return  $time[N] - 1$ 

```

**Computational cost.** See the previous section.

### A.3 Solution 3: BFS-like traversal

```

1: function REACH( $N, M, A[], B[], appear[], disappear[]$ )
2:     for  $i = 0$  to  $M - 1$  do
3:          $edges\_A[i].push\_back(make\_pair(i, B[i]))$ 
4:          $edges\_B[i].push\_back(make\_pair(i, A[i]))$ 
5:     for  $i = 0$  to  $N - 1$  do
6:          $done\_i[i] \leftarrow false$ 
7:          $distance\_i[i] \leftarrow \infty$ 
8:      $reached\_0.push\_back(0)$ 
9:      $distance\_0[0] \leftarrow 0$ 
10:    for  $t = 0$  to  $MAX\_TIME$  do
11:        for all  $v \in reached\_t$  do
12:            if not  $done\_i[v]$  then
13:                for all  $edge \in edges\_v[v]$  do
14:                     $staircase \leftarrow edge.first$ 
15:                     $neighbor \leftarrow edge.second$ 
16:                     $time \leftarrow \max(distance\_i[v], appear[staircase]) + 1$ 
17:                    if not  $done\_i[neighbor]$ 
18:                        and  $distance\_i[v] < disappear[staircase]$ 
19:                        and  $time < distance\_i[neighbor]$  then
20:                             $distance\_i[neighbor] \leftarrow time$ 
21:                             $reached\_time.push\_back(neighbor)$ 
22:                     $done\_i[v] \leftarrow true$ 
23:    return  $(distance\_i[N - 1] = \infty) ? -1 : distance\_i[N - 1]$ 

```

**Computational cost:**  $O(m + MAX\_TIME)$ .

## B Paletta

Paletta ordering<sup>2</sup> is a peculiar ordering technique: given a 3-tuple of elements, paletta takes the central element as pivot and swaps the two elements right before and next to it. To make an example:

$$(3, 2, 1) \xrightarrow{\text{paletta}} (1, 2, 3)$$

We now want to develop an algorithm to order any array through paletta ordering with the minimal number of swaps. You should see as not every array can be ordered (e.g.  $[1, 3, 2]$ ).

**Solution.** We should note that the following properties hold:

1. Every element can be a pivot, but the first and the last one, as they have respectively no elements before and after them.
2. Every element can be swapped as many times as necessary, but only with elements of the same 2-remainder (numbers in even positions can only be swapped with numbers in even positions, the same holds for odd indexes). More formally, if  $n$  is the size of the array  $A$  we want to sort,  $i, j \in [1, n-2]$ ,  $A[i]$  can be swapped with  $A[j]$  if and only if  $i \equiv j \pmod{2}$ .
3. The least number of swaps does not backtrack any element. Formally, let  $k$  be the minimal number of swaps applied to an array, backtracks included. By hypothesis,  $k$  is minimal, but at least  $m$ ,  $m > 0$  backtrack swaps have been operated, therefore we found a  $k' = k - m : k' < k$ , a new minimal number of swaps: contradiction.

Given item 2, we can split our array in two, even and odd numbers, and order them counting the swaps. In our example we'll use *mergesort*, as it runs in  $O(n \log n)$ , does backtrack elements, and is very well-known. Clearly, given an array, a swap happens when an element is pushed back, pulling the one between its new position and the old one ahead: we can map this behaviour in the merge routine of mergesort: the array merged is able to push back elements from its right pointer to the new array, moving them back of  $(m-i) + (j-m)$  positions, where  $m$  is the dimension of the current two sub-arrays to merge. Provided that our edited version of mergesort ran successfully on both the even-index and odd-index, we now need to verify if by merging them we obtain an ordered array. Intuitively, the merged array will start with the first element of the even-index arrays, followed by the first of the odd-index array, followed by the second of the even-index array, and so on. To check for these elements is pretty trivial and can be done in linear time. Follows the pseudo-code for the edited version and SNAKE\_CHECK function:

```
1: function MERGE_WITH_PALETTA(left, right, k):  
2:   ... ▷ merge instructions  
3:   if right > left then  
4:     paletta_count ← paletta_count + 1  
5:     ...  
1: function SNAKE_CHECK  
2:   even, odd ← 0  
3:   for ;even, odd < N; even = even + 1, odd = odd + 1 do  
4:     if a[even] > a[odd] then  
5:       return -1  
6:   return paletta_count
```

**Computational cost:**  $\Theta(n \log n)$ .

---

<sup>2</sup>[http://didawiki.cli.di.unipi.it/lib/exe/fetch.php/magistraleinformatica/alg2/algo2\\_16/paletta.pdf](http://didawiki.cli.di.unipi.it/lib/exe/fetch.php/magistraleinformatica/alg2/algo2_16/paletta.pdf)