

Deterministic and Randomized Quicksort

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Overview

- ➊ Deterministic Quicksort
- ➋ Modify Quicksort to obtain better asymptotic bound
- ➌ Linear-time median algorithm
- ➍ Randomized Quicksort

Deterministic Quicksort

```
Quicksort(A,p,r)
```

```
if p < r then
```

```
    q := Partition(A,p,r); // rearrange A[p..r] in place
```

```
    Quicksort(A, p,q-1);
```

```
    Quicksort(A,p+1,r);
```

Divide-and-Conquer

The design of Quicksort is based on the divide-and-conquer paradigm.

- a) **Divide**: Partition the array $A[p..r]$ into two (possibly empty) subarrays $A[p..q-1]$ and $A[q+1,r]$ such that
 - $A[x] \leq A[q]$ for all x in $[p..q-1]$
 - $A[x] > A[q]$ for all x in $[q+1,r]$
- b) **Conquer**: Recursively sort $A[p..q-1]$ and $A[q+1,r]$
- c) **Combine**: nothing to do here

Partition

2	1	3	4	7	5	6	8
p		i					r

Select pivot (orange element) and rearrange:

- larger elements to the left of the pivot (red)
- elements not exceeding the pivot to the right (yellow)

Partition

Partition(A,p,r)

```
x := A[r]; // select rightmost element as pivot  
i := p-1;  
  
for j = p to r-1 {  
    if A[j] <= x then i := i+1; swap(A[i], A[j]);  
}  
swap(A[i+1],A[r]);  
  
return i+1;
```

Throughout the for loop:

- If $p \leq k \leq i$ then $A[k] \leq x$
- If $i+1 \leq k \leq j-1$ then $A[k] > x$
- If $k=r$, then $A[k] = x$
- $A[j..r-1]$ is unstructured

Partition - Loop - Example

	2	8	7	1	3	5	6	4
i	p,j							r

	2	1	7	8	3	5	6	4
p	i				j			r

	2	8	7	1	3	5	6	4
p,i	j							r

	2	1	3	8	7	5	6	4
p	i				j			r

	2	8	7	1	3	5	6	4
p,i		j						r

	2	1	3	8	7	5	6	4
p		i			j			r

	2	8	7	1	3	5	6	4
p,i			j					r

	2	1	3	8	7	5	6	4
p		i						r

After the loop, the partition routine swaps the leftmost element of the right partition with the pivot element:

	2	1	3	8	7	5	6	4
p		i						r

`swap(A[i+1],A[r])`

	2	1	3	4	7	5	6	8
p		i						r

now recursively sort yellow and red parts.

Worst-Case Partitioning

The worst-case behavior for **quicksort** occurs on an input of length n when partitioning produces just one subproblem with $n-1$ elements and one subproblem with 0 elements.

Therefore the recurrence for the running time $T(n)$ is:

$$T(n) = T(n-1) + T(0) + \Theta(n) = T(n-1) + \Theta(n) = \Theta(n^2)$$

Perhaps we should call this algorithm **pokysort**?

“Better” Quicksort and Linear Median Algorithm

Best-case Partitioning

Best-case partitioning:

If partition produces two subproblems that are roughly of the same size, then the recurrence of the running time is

$$T(n) \leq 2T(n/2) + \theta(n)$$

so that $T(n) = O(n \log n)$

Can we achieve this bound?

Yes, modify the algorithm. Use a linear-time median algorithm to find median, then partition using median as pivot.

Linear Median Algorithm

Let $A[1..n]$ be an array over a totally ordered domain.

- Partition A into groups of 5 and find the median of each group.
[You can do that with 6 comparisons]
- Make an array $U[1..n/5]$ of the medians and find the median m of U by recursively calling the algorithm.
- Partition the array A using the median-of-medians m to find the rank of m in A. If m is of larger rank than the median of A, eliminate all elements $> m$. If m is of smaller rank than the median of A, then eliminate all elements $\leq m$. Repeat the search on the smaller array.

Linear-Time Median Finding

How many elements do we eliminate in each round?

The array U contains $n/5$ elements. Thus, $n/10$ elements of U are larger (smaller) than m , since m is the median of U . Since each element in U is a median itself, there are $3n/10$ elements in A that are larger (smaller) than m .

Therefore, we eliminate $(3/10)n$ elements in each round.

Thus, the time $T(n)$ to find the median is

$$T(n) \leq T(n/5) + T(7n/10) + 6n/5.$$

// median of U , recursive call, and finding medians of groups

Solving the Recurrence

Suppose that $T(n) \leq cn$ (for some c to be determined later)

$$T(n) \leq c(n/5) + c(7n/10) + 6n/5 = c(9n/10) + 6n/5$$

If this is to be $\leq cn$, then we need to have

$$c(9n/10) + 12n/10 \leq cn \text{ or } 12 \leq c$$

Suppose that $T(1) = d$. Then choose $c = \max\{12, d\}$.

An easy proof by induction yields $T(n) \leq cn$.

Goal Achieved?

We can accomplish that quicksort achieves $O(n \log n)$ running time, if we use the linear-time median finding algorithm to select the pivot element.

Unfortunately, the constant in the big Oh expression becomes large, and quicksort loses some of its appeal.

Is there a simpler solution?

Randomized Quicksort

Randomized Quicksort

Randomized-Quicksort(A, p, r)

if $p < r$ then

$q := \text{Randomized-Partition}(A, p, r);$

 Randomized-Quicksort($A, p, q-1$);

 Randomized-Quicksort($A, p+1, r$);

Partition

Randomized-Partition(A, p, r)

$i := \text{Random}(p, r);$

$\text{swap}(A[i], A[r]);$

$\text{Partition}(A, p, r);$

Almost the same as Partition, but now the pivot element is not the rightmost element, but rather an element from $A[p..r]$ that is chosen uniformly at random.

Goal

- The running time of quicksort depends mostly on the number of comparisons performed in all calls to the Randomized-Partition routine.
- Let X denote the random variable counting the number of comparisons in all calls to Randomized-Partition.

Notations

- Let z_i denote the i -th smallest element of $A[1..n]$.
- Thus $A[1..n]$ sorted is $\langle z_1, z_2, \dots, z_n \rangle$.
- Let $Z_{ij} = \{z_i, \dots, z_j\}$ denote the set of elements between z_i and z_j , including these elements.
- $X_{ij} = I\{ z_i \text{ is compared to } z_j \}$.
- Thus, X_{ij} is an indicator random variable for the event that the i -th smallest and the j -th smallest elements of A are compared in an execution of quicksort.

Number of Comparisons

Since each pair of elements is compared at most once by quicksort, the number X of comparisons is given by

$$X = \sum_{i=1}^{n-1} \sum_{j=i+1}^n X_{ij}$$

Therefore, the expected number of comparisons is

$$\mathbb{E}[X] = \sum_{i=1}^{n-1} \sum_{j=i+1}^n \mathbb{E}[X_{ij}] = \sum_{i=1}^{n-1} \sum_{j=i+1}^n \Pr[z_i \text{ is compared to } z_j]$$

When do we compare?

When do we compare z_i to z_j ?

Suppose we pick a pivot element in $Z_{ij} = \{z_i, \dots, z_j\}$.

If $z_i < x < z_j$ then z_i and z_j will land in different partitions and will **never** be compared afterwards.

Therefore, z_i and z_j will be compared if and only if the first element of Z_{ij} to be picked as pivot element is contained in the set $\{z_i, z_j\}$.

Probability of Comparison

$$\begin{aligned} & \Pr[z_i \text{ or } z_j \text{ is the first pivot chosen from } Z_{ij}] \\ &= \Pr[z_i \text{ is the first pivot chosen from } Z_{ij}] \\ &\quad + \Pr[z_j \text{ is the first pivot chosen from } Z_{ij}] \\ &= \frac{1}{j-i+1} + \frac{1}{j-i+1} = \frac{2}{j-i+1} \end{aligned}$$

Expected Number of Comparisons

$$\begin{aligned} \mathbb{E}[X] &= \sum_{i=1}^{n-1} \sum_{j=i+1}^n \frac{2}{j-i+1} \\ &= \sum_{i=1}^{n-1} \sum_{k=1}^{n-i} \frac{2}{k+1} \\ &< \sum_{i=1}^{n-1} \sum_{k=1}^n \frac{2}{k} \\ &= \sum_{i=1}^{n-1} O(\log n) \\ &= O(n \log n) \end{aligned}$$

Conclusion

It follows that the expected running time of Randomized-Quicksort is $O(n \log n)$.

It is unlikely that this algorithm will choose a terribly unbalanced partition each time, so the performance is very good almost all the time.