

CS711008Z Algorithm Design and Analysis

Lecture 5. Basic algorithm design technique: DIVIDE AND CONQUER

Dongbo Bu

Institute of Computing Technology
Chinese Academy of Sciences, Beijing, China

- The basic idea of DIVIDE AND CONQUER technique;
- The first example: MERGESORT
 - Correctness proof by using **loop invariant** technique;
 - Time complexity analysis of recursive algorithm.
- Other examples: COUNTINGINVERSION, CLOSESTPAIR, MULTIPLICATION, FFT;
- Combining with randomization: QUICKSORT, QUICKSELECT, BFPRT and FLOYDRIVEST algorithm for SELECTION problem;
- Remarks:
 - 1 DIVIDE AND CONQUER could serve to reduce the running time though **the brute-force algorithm is already polynomial-time**, say the $O(n^2)$ brute-force algorithm versus $O(n \log n)$ divide and conquer algorithm for the CLOSESTPAIR problem.
 - 2 This technique is especially powerful when **combined with randomization technique**.

The general DIVIDE AND CONQUER paradigm

- Basic idea: Many problems are recursive in structure, i.e., to solve a given problem, they call themselves several times to deal with closely related **sub-problems**. These sub-problems have the same form to the original problem but a smaller size.
- Three steps of the DIVIDE AND CONQUER paradigm:
 - ① **Divide** a problem into a number of **independent sub-problems**;
 - ② **Conquer** the subproblems by solving them recursively;
 - ③ **Combine** the solutions to the subproblems into the solution to the original problem.

DIVIDE AND CONQUER technique

- To see whether the **DIVIDE AND CONQUER** technique applies on a given problem, we need to examine both **input** and **output** of the problem description.
 - Examine the **input** part to determine how to decompose the problem into subproblems of same structure but smaller size: It is relatively easy to decompose a problem into subproblems if the input part is related to the following data structures:
 - An **array** with n elements;
 - A **matrix**;
 - A **set** of n elements;
 - A **tree**;
 - A **directed acyclic graph**;
 - A **general graph**.
 - Examine the **output** part to determine how to construct the solution to the original problem using the solutions to its subproblems.

SORT problem: to sort an **array** of n integers

SORT problem

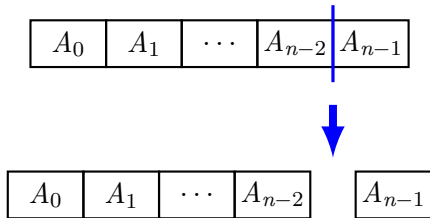
INPUT: An array of n integers, denoted as $A[0..n-1]$;

OUTPUT: The elements of A in increasing order.

- An array can be divided into smaller ones based on **indices** or **values** of elements.

Divide strategy 1 based on indices of elements

- **Divide array $A[0..n-1]$ into a $n-1$ -length array $A[0..n-2]$ and a single element:** $A[0..n-2]$ has the same form to $A[0..n-1]$ but smaller size; thus, sorting $A[0..n-2]$ constructs a subproblem of the original problem. The **DIVIDE AND CONQUER** strategy might apply if we can sort $A[0..n-1]$ using the sorted $A[0..n-2]$.



Sort $A[0..n-1]$ using the sorted $A[0..n-2]$

- Basic idea: To sort $A[0..n-1]$, it suffices to put $A[n-1]$ in its correct position among the sorted $A[0..n-2]$, which can be achieved through comparing $A[n-1]$ with the elements in $A[0..n-2]$.

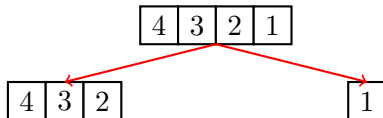
INSERTSORT(A, k)

```
1: if  $k \leq 1$  then  
2:   return ;  
3: end if  
4: INSERTSORT( $A, k-1$ );  
5:  $key = A[k]$ ;  
6:  $i = k-1$ ;  
7: while  $i \geq 0$  and  $A[i] > key$  do  
8:    $A[i+1] = A[i]$ ;  
9:    $i = i-1$ ;  
10: end while  
11:  $A[i+1] = key$ ;
```


An example

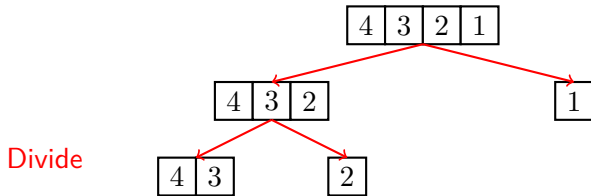
4	3	2	1
---	---	---	---

An example

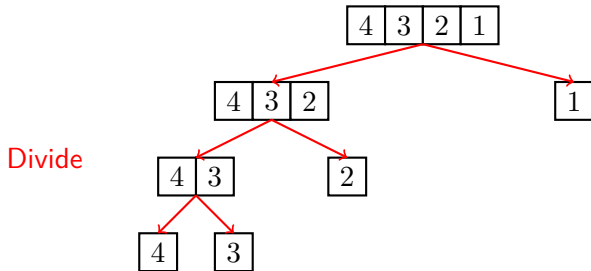


Divide

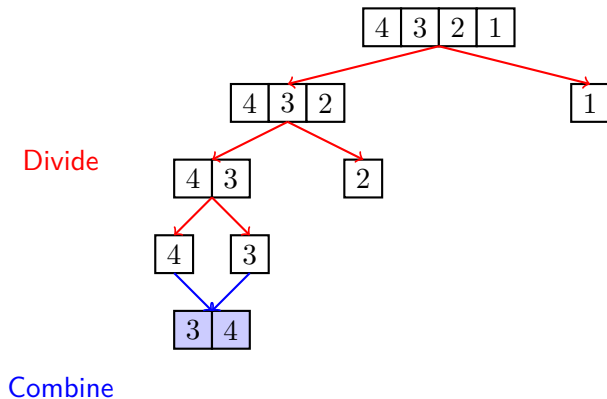
An example



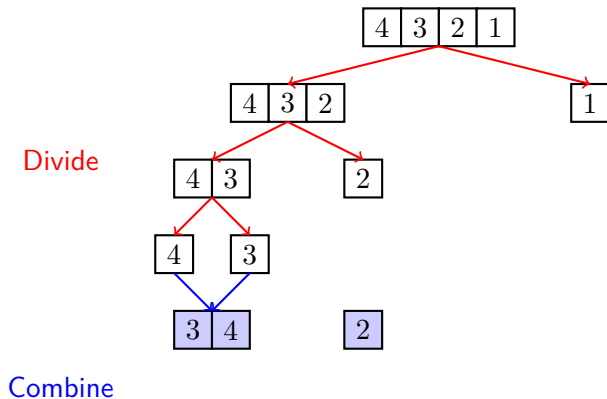
An example



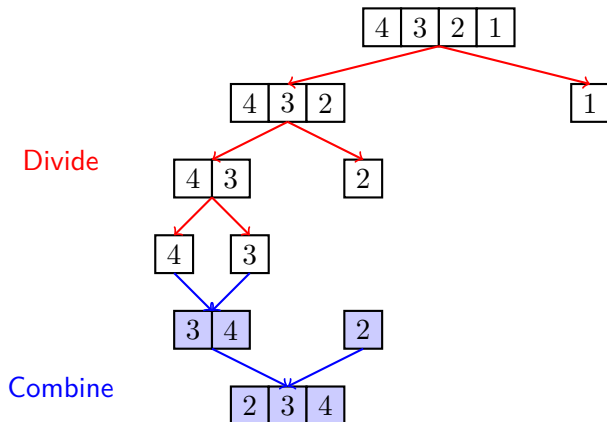
An example



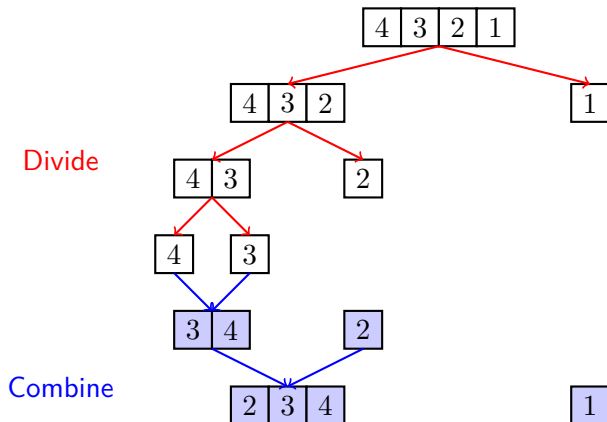
An example



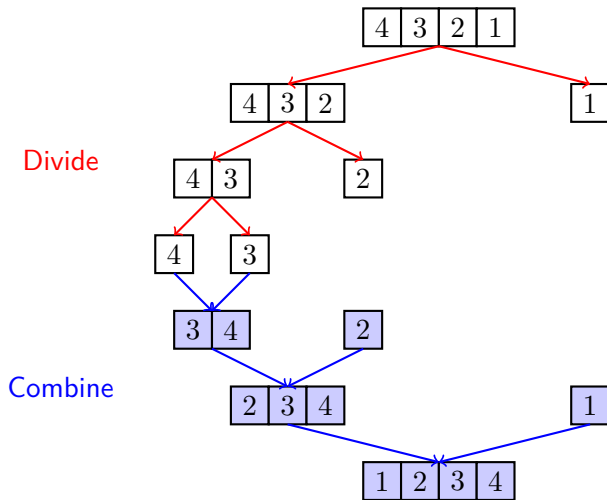
An example



An example

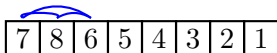
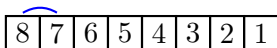


An example

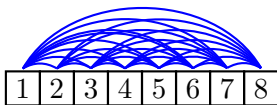


Analysis of INSERTSORT algorithm

- Worst case: elements in $A[0..n-1]$ are in decreasing order.
- Time complexity: $T(n) = T(n-1) + O(n) = O(n^2)$. The subproblems decrease **slowly in size** (linearly here, reducing by only one element each time); thus the sum of linear steps yields quadratic overall time.



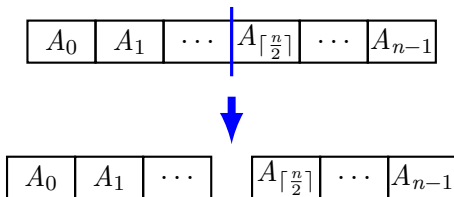
⋮



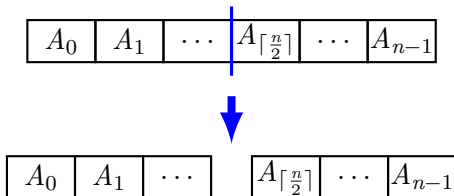
INSERTSORT: 28 ops

Divide strategy 2 based on indices of elements

- **Divide the array $A[0..n-1]$ into two arrays $A[0..\lceil \frac{n}{2} \rceil - 1]$ and $A[\lceil \frac{n}{2} \rceil..n-1]$:** Both $A[0..\lceil \frac{n}{2} \rceil - 1]$ and $A[\lceil \frac{n}{2} \rceil..n-1]$ have same form to $A[0..n-1]$ but smaller size; thus, sorting $A[0..\lceil \frac{n}{2} \rceil - 1]$ and $A[\lceil \frac{n}{2} \rceil..n-1]$ construct two subproblem of the original problem. The **DIVIDE AND CONQUER** technique might apply if we can sort $A[0..n-1]$ using the sorted $A[0..\lceil \frac{n}{2} \rceil - 1]$ and the sorted $A[\lceil \frac{n}{2} \rceil..n-1]$.



MERGESORT algorithm [J. von Neumann, 1945, 1948]



MERGESORT(A, l, r)

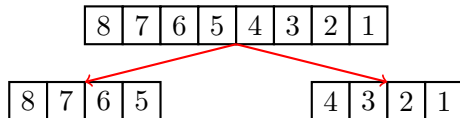
- 1: //Sort elements in $A[l..r]$
- 2: **if** $l < r$ **then**
- 3: $m = (l + r)/2$; // m denotes the middle point
- 4: MERGESORT(A, l, m);
- 5: MERGESORT($A, m + 1, r$);
- 6: MERGE(A, l, m, r); //Combining the sorted arrays
- 7: **end if**

- Sort the entire array: MERGESORT($A, 0, n - 1$)

An example

8	7	6	5	4	3	2	1
---	---	---	---	---	---	---	---

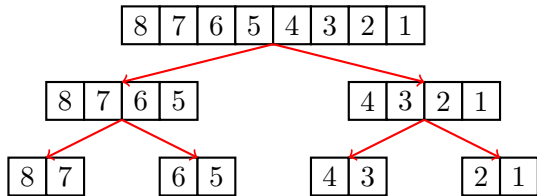
An example



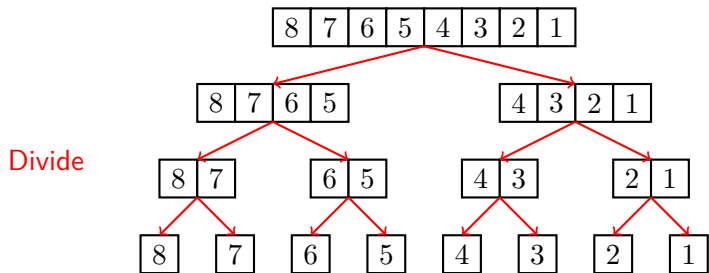
Divide

An example

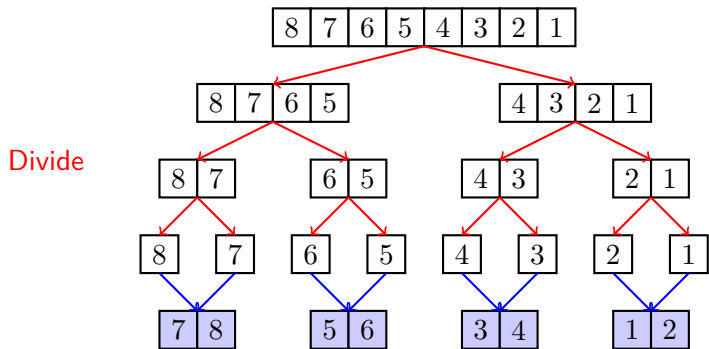
Divide



An example

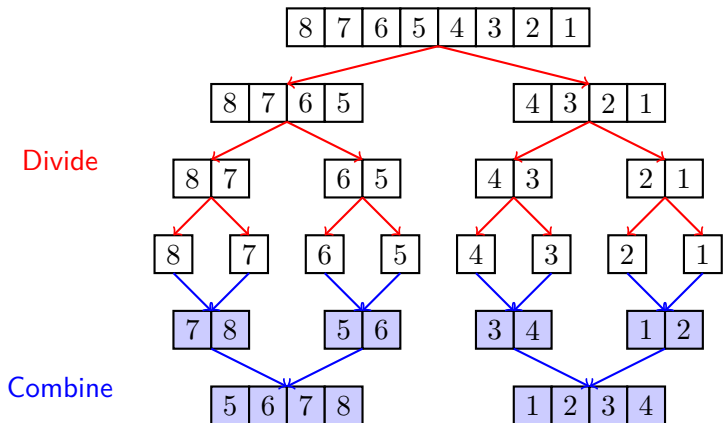


An example

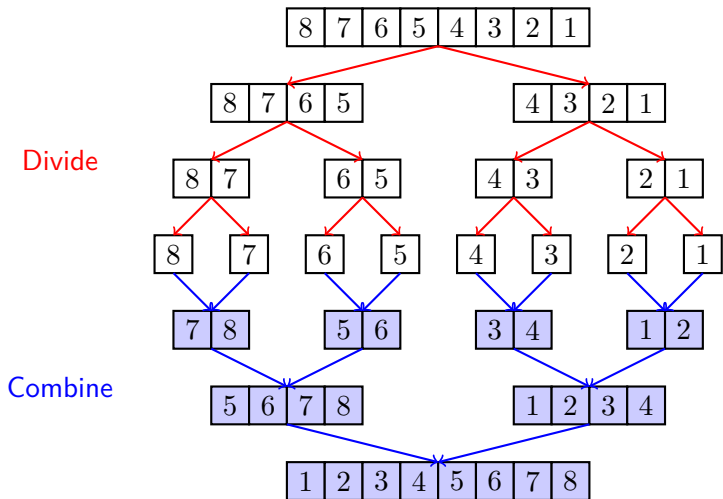


Combine

An example



An example

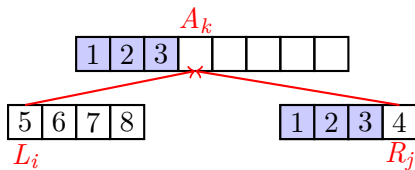


MERGESORT algorithm: how to combine?

MERGE (A, l, m, r)

```
1: //Merge  $A[l..m]$  (denoted as  $L$ ) and  $A[m+1..r]$  (denoted as  $R$ ).
2:  $i = 0; j = 0;$ 
3: for  $k = l$  to  $r$  do
4:   if  $L[i] < R[j]$  then
5:      $A[k] = L[i];$ 
6:      $i++;$ 
7:   if all elements in  $L$  have been copied then
8:     Copy the remainder elements from  $R$  into  $A$ ;
9:     break;
10:  end if
11: else
12:    $A[k] = R[j];$ 
13:    $j++;$ 
14:   if all elements in  $R$  have been copied then
15:     Copy the remainder elements from  $L$  into  $A$ ;
16:     break;
17:   end if
18: end if
19: end for
```

MERGE algorithm



(see a demo)

Correctness of MERGESORT algorithm

Correctness of **Merge** procedure: **loop-invariant** technique [R. W. Floyd, 1967]

Loop invariant: (similar to **mathematical induction** proof technique)

- 1 At the start of each iteration of the **for** loop, $A[l..k-1]$ contains the $k-l$ smallest elements of $L[1..n_1+1]$ and $R[1..n_2+1]$, in sorted order.
- 2 $L[i]$ and $R[j]$ are the smallest elements of their array that have not been copied to A .

Proof.

- Initialization: $k = l$. Loop invariant holds since $A[l..k-1]$ is empty.
- Maintenance: Suppose $L[i] < R[j]$, and $A[l..k-1]$ holds the $k-l$ smallest elements. After copying $L[i]$ into $A[k]$, $A[l..k]$ will hold the $k-l+1$ smallest elements.



Correctness of **Merge** procedure: **loop-invariant** technique [R. W. Floyd, 1967]

- Since the loop invariant holds initially, and is maintained during the **for** loop, thus it should hold when the algorithm terminates.
- Termination: At termination, $k = r + 1$. By loop invariant, $A[l..k - 1]$, i.e. $A[l..r]$ must contain $r - l + 1$ smallest elements, in sorted order.

Time-complexity of MERGESORT algorithm

Time-complexity of MERGE algorithm

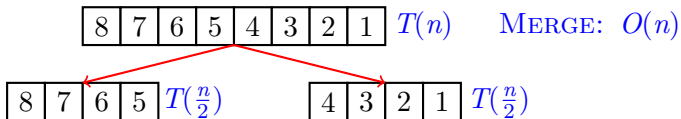
MERGE (A, l, m, r)

```
1: //Merge  $A[l..m]$  (denoted as  $L$ ) and  $A[m + 1..r]$  (denoted as  $R$ ).
2:  $i = 0; j = 0;$ 
3: for  $k = l$  to  $r$  do
4:   if  $L[i] < R[j]$  then
5:      $A[k] = L[i];$ 
6:      $i++;$ 
7:   if all elements in  $L$  have been copied then
8:     Copy the remainder elements from  $R$  into  $A;$ 
9:     break;
10:  end if
11: else
12:    $A[k] = R[j];$ 
13:    $j++;$ 
14:  if all elements in  $R$  have been copied then
15:    Copy the remainder elements from  $L$  into  $A;$ 
16:    break;
17:  end if
18: end if
19: end for
```

Time complexity: $O(n)$.

Time-complexity of MERGESORT algorithm

- Let $T(n)$ denote the running time of MERGESORT on an array of size n . As comparison of elements dominates the algorithm, we use the number of comparisons as $T(n)$.



- We have the following recursion:

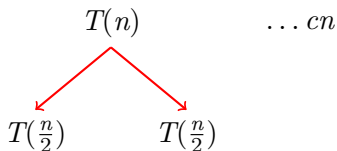
$$T(n) = \begin{cases} 1 & \text{if } n \leq 2 \\ T(\frac{n}{2}) + T(\frac{n}{2}) + O(n) & \text{otherwise} \end{cases} \quad (1)$$

- Note that the subproblems decrease **exponentially in size**, which is much faster than the linearly decrease in INSERTSORT.

- Ways to analyse a recursion:
 - ① **Unrolling the recurrence:** unrolling a few levels to find a pattern, and then sum over all levels;
 - ② **Guess and substitution:** guess the solution, substitute it into the recurrence relation, and check whether it works.
 - ③ **Master theorem**

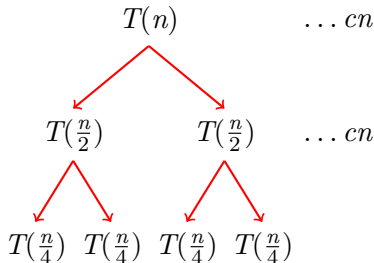
Analysis technique 1: Unrolling the recurrence

- We have $T(n) = 2T(\frac{n}{2}) + O(n) \leq 2T(\frac{n}{2}) + cn$ for a constant c . Let unrolling a few levels to find a pattern, and then sum over all levels.



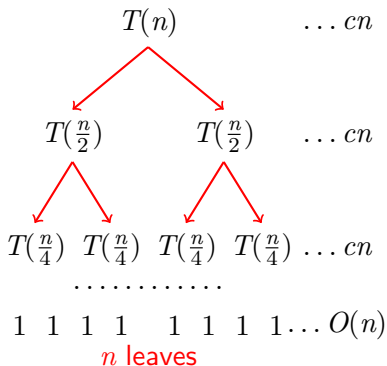
Analysis technique 1: Unrolling the recurrence

- We have $T(n) = 2T(\frac{n}{2}) + O(n) \leq 2T(\frac{n}{2}) + cn$ for a constant c . Let unrolling a few levels to find a pattern, and then sum over all levels.



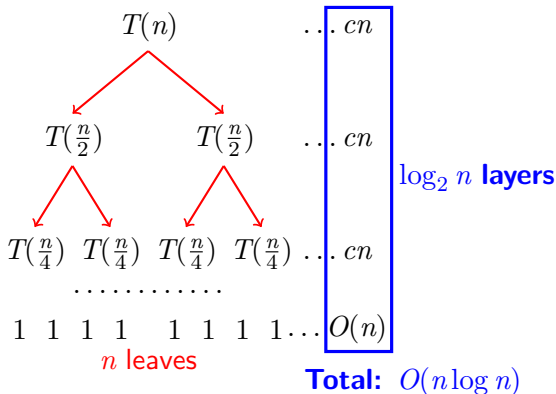
Analysis technique 1: Unrolling the recurrence

- We have $T(n) = 2T(\frac{n}{2}) + O(n) \leq 2T(\frac{n}{2}) + cn$ for a constant c . Let unrolling a few levels to find a pattern, and then sum over all levels.



Analysis technique 1: Unrolling the recurrence

- We have $T(n) = 2T(\frac{n}{2}) + O(n) \leq 2T(\frac{n}{2}) + cn$ for a constant c . Let unrolling a few levels to find a pattern, and then sum over all levels.



Analysis technique 2: Guess and substitution

- Guess and substitution: guess a solution, substitute it into the recurrence relation, and justify that it works.
- Guess: $T(n) \leq cn \log_2 n$.
- Verification:
 - Case $n = 2$: $T(2) = 1 \leq cn \log_2 n$;
 - Case $n > 2$: Suppose $T(m) \leq cm \log_2 m$ holds for all $m \leq n$.
We have

$$\begin{aligned}T(n) &= 2T\left(\frac{n}{2}\right) + cn \\&\leq 2c\frac{n}{2} \log_2\left(\frac{n}{2}\right) + cn \\&= 2c\frac{n}{2} \log_2 n - 2c\frac{n}{2} + cn \\&= cn \log_2 n\end{aligned}$$

Analysis technique 2: a weaker version

- Guess and substitution: one guesses the overall form of the solution without pinning down the constants and parameters.
- A weaker guess: $T(n) = O(n \log n)$. Rewritten as $T(n) \leq kn \log_b n$, where k, b **will be determined later**.

$$\begin{aligned} T(n) &\leq 2T\left(\frac{n}{2}\right) + cn \\ &\leq 2k\frac{n}{2} \log_b\left(\frac{n}{2}\right) + cn \quad (\text{set } b = 2 \text{ for simplification}) \\ &= 2k\frac{n}{2} \log_2 n - 2k\frac{n}{2} + cn \\ &= kn \log_2 n - kn + cn \quad (\text{set } k = c \text{ for simplification}) \\ &= cn \log_2 n \end{aligned}$$

Theorem

Let $T(n)$ be defined by $T(n) = aT(\frac{n}{b}) + O(n^d)$ for $a > 1$, $b > 1$ and $d > 0$, then $T(n)$ can be bounded by:

- 1 If $d < \log_b a$, then $T(n) = O(n^{\log_b a})$;
- 2 If $d = \log_b a$, then $T(n) = O(n^{\log_b a} \log n)$;
- 3 If $d > \log_b a$, then $T(n) = O(n^d)$.

- Intuition: the ratio of cost between neighbouring layers is $\frac{a}{b^d}$.

Proof.

$$\begin{aligned}T(n) &= aT\left(\frac{n}{b}\right) + O(n^d) \\&\leq aT\left(\frac{n}{b}\right) + cn^d \\&\leq a\left(aT\left(\frac{n}{b^2}\right) + c\left(\frac{n}{b}\right)^d\right) + cn^d \\&\leq \dots\dots\dots \\&\leq cn^d\left(1 + \frac{a}{b^d} + \left(\frac{a}{b^d}\right)^2 + \dots + \left(\frac{a}{b^d}\right)^{\log_b n - 1}\right) + a^{\log_b n} \\&= \begin{cases} O(n^{\log_b a}) & \text{if } d < \log_b a \\ O(n^{\log_b a} \log n) & \text{if } d = \log_b a \\ O(n^d) & \text{if } d > \log_b a \end{cases}\end{aligned}$$

Here $c > 0$ represents a constant. □

Master theorem: examples

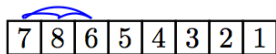
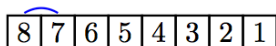
- Example 1: $T(n) = 3T(\frac{n}{2}) + O(n)$

$$T(n) = O(n^{\log_2 3}) = O(n^{1.585})$$

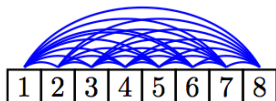
- Example 2: $T(n) = 2T(\frac{n}{2}) + O(n^2)$

$$T(n) = O(n^2)$$

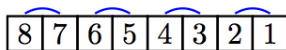
Question: from $O(n^2)$ to $O(n \log n)$, what did we save?



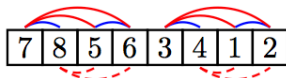
⋮



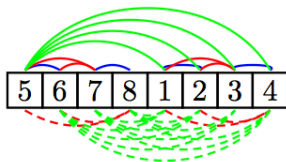
INSERTSORT: 28 ops



MERGESORT step 1: 4 ops



MERGESORT step 2: 4 ops, save: 4



MERGESORT step 3: 4 ops, save: 12

COUNTINGINVERSION: to count inversions in an **array** of n integers

COUNTINGINVERSION problem

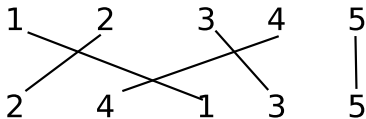
Practical problems:

- 1 To identify two users with similar preference, i.e. ranking books, movies, etc.

COUNTINGINVERSION problem

INPUT: An array $A[0..n-1]$ with n distinct numbers;

OUTPUT: the number of **inversions**. A pair of indices i and j constitutes an inversion if $i < j$ but $A[i] > A[j]$.



Application 1: Genome comparison

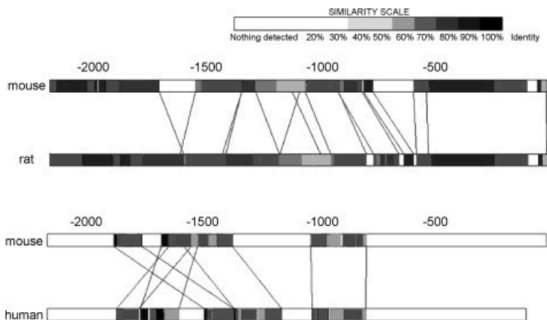


Figure 1: Sequence comparison of the 5' flanking regions of mouse, rat and human ER β .

Reference: In vivo function of the 5' flanking region of mouse estrogen receptor β gene, The Journal of Steroid Biochemistry and Molecular Biology Volume 105, Issues 1-5, June-July 2007, pages 57-62.

Application 2: A measure of bivariate association

- Motivation: how to measure the association between two genes when given expression levels across n time points?
- Existing measures:
 - Linear relationship: Pearson's CC (most widely used, but sensitive to outliers)
 - Monotonic relationship: Spearman, Kendall's correlation
 - General statistical dependence: Renyi correlation, mutual information, maximal information coefficient
- A novel measure:

$$W_1 = \sum_{i=1}^{n-k+1} (I_i^+ + I_i^-)$$

Here, I_i^+ is 1 if $X_{[i,..,i+k-1]}$ and $Y_{[i,..,i+k-1]}$ has the same order and 0 otherwise, while I_i^- is 1 if $X_{[i,..,i+k-1]}$ and $-Y_{[i,..,i+k-1]}$ has the same order and 0 otherwise.

- Advantage: the association may exist across a subset of samples. For example,

X : 1 3 4 2 5

Y : 1 4 5 2 3

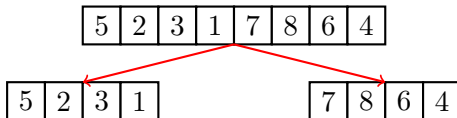
$W_1 = 2$ when $k = 3$. Much better than Pearson CC, et al.

- Solution: index pairs. The possible solution space has a size of $O(n^2)$.
- Brute-force: $O(n^2)$ (Examining all index pairs (i, j)).
- Can we design a better algorithm?

COUNTINGINVERSION problem

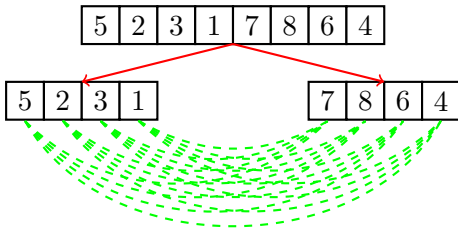
- DIVIDE AND CONQUER technique:

- 1 **Divide:** Divide A into two arrays $A[0..\lceil \frac{n}{2} \rceil - 1]$ and $A[\lceil \frac{n}{2} \rceil .. n - 1]$; thus counting inversions within $A[0..\lceil \frac{n}{2} \rceil - 1]$ and $A[\lceil \frac{n}{2} \rceil .. n - 1]$ constitutes two subproblems.
- 2 **Conquer:** Counting inversions within each half by calling COUNTINGINVERSION itself.



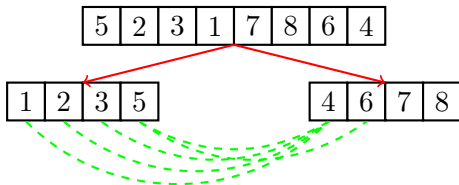
Combine strategy 1

- **Combine:** How to count the inversions (i, j) with $A[i]$ and $A[j]$ from different halves?
- If the two halves $A[0..\lceil \frac{n}{2} \rceil - 1]$ and $A[\lceil \frac{n}{2} \rceil .. n - 1]$ have no special structure, we have to examine all possible index pairs $i \in [0, \lceil \frac{n}{2} \rceil - 1]$, $j \in [\lceil \frac{n}{2} \rceil, n - 1]$ to count such inversions, which costs $\frac{n^2}{4}$ time.
- Thus, $T(n) = 2T(\frac{n}{2}) + \frac{n^2}{4} = O(n^2)$.



Combine strategy 2

- **Combine:** How to count the inversions (i, j) with $A[i]$ and $A[j]$ from different halves?
- If the two halves are unstructured, it would be inefficient to count inversions. Thus, we need to introduce some structures into $A[0..\lceil \frac{n}{2} \rceil - 1]$ and $A[\lceil \frac{n}{2} \rceil .. n - 1]$.
- Note that it is relatively easy to count such inversions if **elements in both halves are in increasing order.**



(See a demo)

SORT-AND-COUNT algorithm

SORT-AND-COUNT(A)

- 1: Divide A into two sub-sequences L and R ;
- 2: $(RC_L, L) = \text{SORT-AND-COUNT}(L)$;
- 3: $(RC_R, R) = \text{SORT-AND-COUNT}(R)$;
- 4: $(C, A) = \text{MERGE-AND-COUNT}(L, R)$;
- 5: **return** $(RC = RC_L + RC_R + C, A)$;

Time complexity: $T(n) = 2T(\frac{n}{2}) + O(n) = O(n \log n)$.

MERGE-AND-COUNT algorithm

MERGE-AND-COUNT (L, R)

```
1:  $RC = 0; i = 0; j = 0;$ 
2: for  $k = 0$  to  $\|L\| + \|R\| - 1$  do
3:   if  $L[i] > R[j]$  then
4:      $A[k] = R[j];$ 
5:      $j++;$ 
6:      $RC += (\|L\| - i);$ 
7:     if all elements in  $R$  have been copied then
8:       Copy the remainder elements from  $L$  into  $A$ ;
9:       break;
10:    end if
11:  else
12:     $A[k] = L[i];$ 
13:     $i++;$ 
14:    if all elements in  $L$  have been copied then
15:      Copy the remainder elements from  $R$  into  $A$ ;
16:      break;
17:    end if
18:  end if
19: end for
```


QUICKSORT algorithm: divide based on **value of elements**



Figure 2: Sir Charles Antony Richard Hoare, 2011

QUICKSORT: divide based on value of a randomly-selected element

QUICKSORT(A)

```
1:  $S_- = \{\}; S_+ = \{\};$   
2: Choose a pivot  $A[j]$  uniformly at random;  
3: for  $i = 0$  to  $n - 1$  do  
4:   Put  $A[i]$  in  $S_-$  if  $A[i] < A[j]$ ;  
5:   Put  $A[i]$  in  $S_+$  if  $A[i] \geq A[j]$ ;  
6: end for  
7: QUICKSORT( $S_+$ );  
8: QUICKSORT( $S_-$ );  
9: Output  $S_-$ , then  $A[j]$ , then  $S_+$ ;
```

- The randomization operation makes this algorithm **simple** (relative to MERGESORT algorithm) but **efficient**.
- However, the randomization also makes it difficult to analyze time-complexity: When dividing based on indices, it is easy to divide into two halves with equal size; in contrast, we divide based on value of a randomly-selected pivot and thus we cannot guarantee that each sub-problem has exactly $\frac{n}{2}$ elements.

Various cases of the execution of QUICKSORT algorithm

- **Worst case:** selecting the smallest/largest element at each iteration. The subproblems decrease **linearly** in size.

$$T(n) = T(n-1) + O(n) = O(n^2)$$

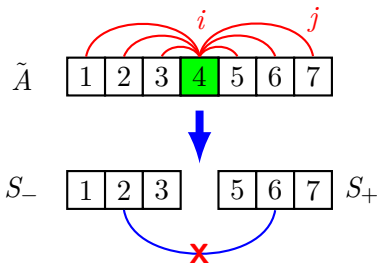
- **Best case:** select the median exactly at each iteration. The subproblems decrease **exponentially** in size.

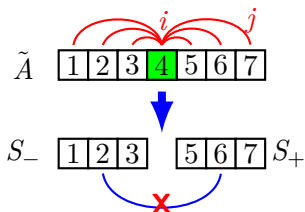
$$T(n) = 2T\left(\frac{n}{2}\right) + O(n) = O(n \log n)$$

- **Most cases:** instead of selecting the median exactly, we can select a **nearly-central pivot** with high probability. We claim that the expected running time is still

$$T(n) = O(n \log n).$$

- Let X denote the number of comparisons performed in line 4 and 5. After expanding all recursive calls, it is obvious that the running time of QUICKSORT is $O(n + X)$. Our objective is to calculate $E[X]$.
- For simplicity, we represent each element using its index in the sorted array, denoted as \tilde{A} . We have two key observations:
- Observation 1:** Any two elements $\tilde{A}[i]$ and $\tilde{A}[j]$ are compared at most once.





- Define index variable

$$X_{ij} = \begin{cases} 1 & \text{if } \tilde{A}[i] \text{ is compared with } \tilde{A}[j] \\ 0 & \text{otherwise} \end{cases}$$

- Thus $X = \sum_{i=0}^{n-1} \sum_{j=i+1}^{n-1} X_{ij}$.

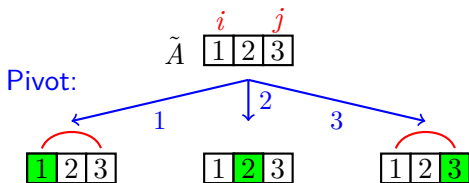
$$\begin{aligned} E[X] &= E\left[\sum_{i=0}^{n-1} \sum_{j=i+1}^{n-1} X_{ij}\right] \\ &= \sum_{i=0}^{n-1} \sum_{j=i+1}^{n-1} E[X_{ij}] \\ &= \sum_{i=0}^{n-1} \sum_{j=i+1}^{n-1} \Pr(\tilde{A}[i] \text{ is compared with } \tilde{A}[j]) \end{aligned}$$

- **Observation 2:** $\tilde{A}[i]$ and $\tilde{A}[j]$ are compared iff either $\tilde{A}[i]$ or $\tilde{A}[j]$ is selected as pivot when processing elements containing $\tilde{A}[i..j]$.
- We claim $\Pr(\tilde{A}[i] \text{ is compared with } \tilde{A}[j]) = \frac{2}{j-i+1}$. (Why?)
- Then

$$\begin{aligned} E[X] &= \sum_{i=0}^{n-1} \sum_{j=i+1}^{n-1} \Pr(\tilde{A}[i] \text{ is compared with } \tilde{A}[j]) \\ &= \sum_{i=0}^{n-1} \sum_{j=i+1}^{n-1} \frac{2}{j-i+1} \\ &= \sum_{i=0}^{n-1} \sum_{k=1}^{n-i-1} \frac{2}{k+1} \\ &\leq \sum_{i=0}^{n-1} \sum_{k=1}^{n-1} \frac{2}{k+1} \\ &= O(n \log n) \end{aligned}$$

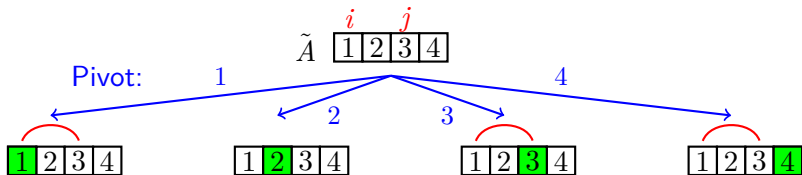
- Here k is defined as $k = j - i$.

Why $\Pr(\tilde{A}[i] \text{ is compared with } \tilde{A}[j]) = \frac{2}{j-i+1}$?



- Let's examine a simple example first: For an array with only 3 elements, each element will be selected as pivot with equal probability $\frac{1}{3}$.
- In two out of the three cases, $\tilde{A}[i]$ is compared with $\tilde{A}[j]$. Hence, $\Pr(\tilde{A}[i] \text{ is compared with } \tilde{A}[j]) = \frac{2}{3}$

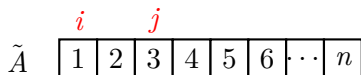
Why $\Pr(\tilde{A}[i] \text{ is compared with } \tilde{A}[j]) = \frac{2}{j-i+1}$? cont'd



- Let's consider a larger array with 4 elements.
- Each element will be selected as pivot with equal probability $\frac{1}{4}$: the selection of $\tilde{A}[i]$ or $\tilde{A}[j]$ as pivot will lead to an immediate comparison of $\tilde{A}[i]$ and $\tilde{A}[j]$. In contrast, the selection of $\tilde{A}[3]$ as pivot produces a smaller problem, where $\tilde{A}[i]$ will be compared with $\tilde{A}[j]$ with probability $\frac{2}{3}$ by induction. Hence,

$$\begin{aligned}\Pr(\tilde{A}[i] \text{ is compared with } \tilde{A}[j]) &= \frac{1}{4} + 0 + \frac{1}{4} + \frac{1}{4} \times \frac{2}{3} \\ &= \frac{3}{4} \times \frac{2}{3} + \frac{1}{4} \times \frac{2}{3} \\ &= \frac{2}{3}\end{aligned}$$

Why $\Pr(\tilde{A}[i] \text{ is compared with } \tilde{A}[j]) = \frac{2}{j-i+1}$? cont'd



- Now let's extend these observations to general case that A has n elements. By induction over the size of A , we can calculate the probability as:

$$\begin{aligned}\Pr(\tilde{A}[i] \text{ is compared with } \tilde{A}[j]) &= \frac{1}{n} + \frac{1}{n} + \frac{n-(j-i+1)}{n} \times \frac{2}{j-i+1} \\ &= \left(\frac{j-i+1}{n} + \frac{n-(j-i+1)}{n} \right) \times \frac{2}{j-i+1} \\ &= \frac{2}{j-i+1}\end{aligned}$$

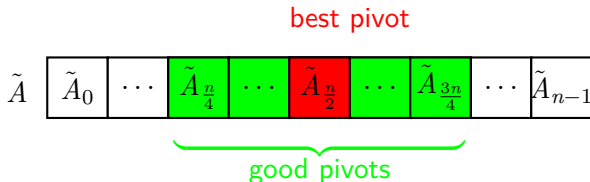
MODIFIED QUICKSORT: easier to analyze

MODIFIEDQUICKSORT(A)

```
1: while TRUE do
2:   Choose a pivot  $A[j]$  uniformly at random;
3:    $S_- = \{\}$ ;  $S_+ = \{\}$ ;
4:   for  $i = 0$  to  $n - 1$  do
5:     Put  $A[i]$  in  $S_-$  if  $A[i] < A[j]$ ;
6:     Put  $A[i]$  in  $S_+$  if  $A[i] \geq A[j]$ ;
7:   end for
8:   if  $\|S_+\| \geq \frac{n}{4}$  and  $\|S_-\| \geq \frac{n}{4}$  then
9:     break; //  $A$  fixed proportion of elements fall both below and
       above the pivot;
10:  end if
11: end while
12: MODIFIEDQUICKSORT( $S_+$ );
13: MODIFIEDQUICKSORT( $S_-$ );
14: Output  $S_-$ , then  $A[j]$ , and finally  $S_+$ ;
```

- MODIFIEDQUICKSORT works when all items are distinct.
However, it is slower than the original version since it doesn't run when the pivot is "off-center".

MODIFIED QUICKSORT: analysis



- It is easy to obtain a **nearly central pivot**:
 - $\Pr(\text{select the centroid as pivot}) = \frac{1}{n}$
 - $\Pr(\text{select a nearly central element as pivot}) = \frac{1}{2}$
 - Thus $E(\#WHILE) = 2$, i.e., the expected time of finding a nearly central pivot is $2n$.
- **Nearly central pivot** is good:
 - An element is a **good pivot** if a fixed proportion of elements fall both below and above it, thus making subproblems decrease **exponentially** in size.
 - Specifically, the recursion tree has a depth of $O(\log_{\frac{4}{3}} n)$, and $O(n)$ work is needed at each level, hence $T(n) = O(n \log_{\frac{4}{3}} n)$.

Lomuto's in-place algorithm

QUICKSORT(A, l, r)

1: **if** $l < r$ **then**

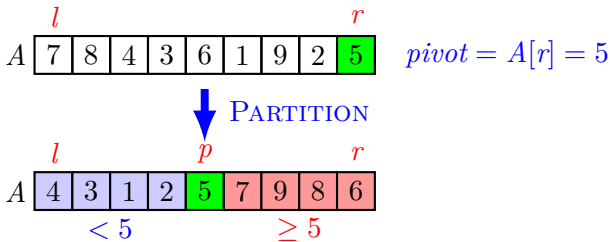
2: $p = \text{PARTITION}(A, l, r)$ //Use $A[r]$ as pivot;

3: QUICKSORT($A, l, p - 1$);

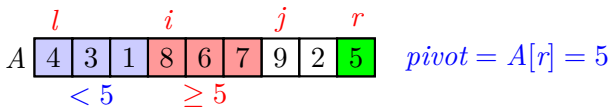
4: QUICKSORT($A, p + 1, r$);

5: **end if**

- Sort the entire array: QUICKSORT($A, 0, n - 1$).



Lomuto's PARTITION procedure



- Basic idea: Swap the elements (in $A[l..j-1]$) to make elements in $A[l..i-1] < pivot$ and elements in $A[i..j-1] \geq pivot$.

PARTITION(A, l, r)

- 1: $pivot = A[r]; i = l;$
- 2: **for** $j = l$ to $r - 1$ **do**
- 3: **if** $A[j] < pivot$ **then**
- 4: Swap $A[i]$ with $A[j];$
- 5: $i++;$
- 6: **end if**
- 7: **end for**
- 8: Swap $A[i]$ with $A[r];$ // Put pivot in its correct position
- 9: **return** $i;$

Hoare's in-place algorithm [1961]

QUICKSORT(A, l, r)

1: **if** $l < r$ **then**

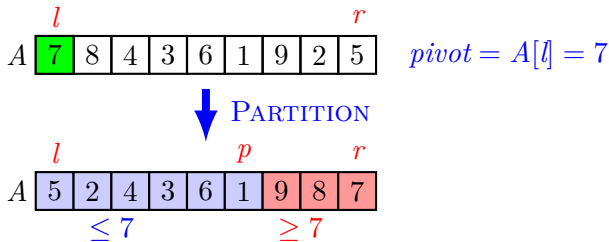
2: $p = \text{PARTITION}(A, l, r)$ //Use $A[l]$ as pivot;

3: QUICKSORT(A, l, p); //Reason: $A[p]$ might not be at its correct position

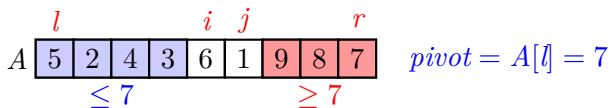
4: QUICKSORT($A, p + 1, r$);

5: **end if**

- Sort the entire array: QUICKSORT($A, 0, n - 1$).



Hoares' PARTITION procedure



- Basic idea: Keep the elements in $A[l..i-1] \leq pivot$ and the elements in $A[j+1..r] \geq pivot$.

PARTITION(A, l, r)

```
1:  $i = l - 1; j = r + 1; pivot = A[l];$ 
2: while TRUE do
3:   repeat
4:      $j = j - 1$ ; //From right to left, find the first element  $\leq pivot$ 
5:   until  $A[j] \leq pivot$  or  $j == l$ ;
6:   repeat
7:      $i = i + 1$ ; //From left to right, find the first element  $\geq pivot$ 
8:   until  $A[i] \geq pivot$  or  $i == r$ ;
9:   if  $j \leq i$  then
10:    return  $j$ ;
11:  end if
12:  Swap  $A[i]$  with  $A[j]$ ;
13: end while
```


Comparison with MERGESORT [Hoare, 1961]

NUMBER OF ITEMS	MERGE SORT	QUICKSORT
500	2 min 8 sec	1 min 21 sec
1,000	4 min 48 sec	3 min 8 sec
1,500	8 min 15 sec*	5 min 6 sec
2,000	11 min 0 sec*	6 min 47 sec

* These figures were computed by formula, since they cannot be achieved on the 405 owing to limited store size.

- Note: The preceding QUICKSORT algorithm works well for lists with **distinct elements** but exhibits poor performance when the input list contains many **repeated elements**. To solve this problem, an alternative PARTITION algorithm was proposed to divide the list into three parts: elements less than pivot, elements equal to pivot, and elements greater than pivot. Only the less-than and greater-than pivot partitions need to be recursively sorted.

Extension: sorting on dynamic data

- When the data changes gradually, the goal of a sorting algorithm is to sort the data at each time step, under the constraint that it only has limited access to the data each time.
- As the data is constantly changing and the algorithm might be unaware of these changes, it cannot be expected to always output the exact right solution; we are interested in algorithms that guarantee to output an approximate solution.
- In 2011, Eli Upfal et al. proposed an algorithm to sort dynamic data.
- In 2017, Liu and Huang proposed an efficient algorithm to determine top k elements of dynamic data.

SELECTION problem: to select the k -th smallest items in **an array**

INPUT:

An array $A = [A_0, A_1, \dots, A_{n-1}]$, and a number $k < n$;

OUTPUT:

The k -th smallest item in general case (or the median of A as a special case).

- Things will be easy when k is very small, say $k = 1, 2$.
However, identification of the median is not that easy.
- The k -th smallest element could be readily determined after sorting A , which takes $O(n \log n)$ time.
- In contrast, when using DIVIDE AND CONQUER technique, it is possible to develop a faster algorithm, say the deterministic linear algorithm ($16n$ comparisons) by Blum et al.

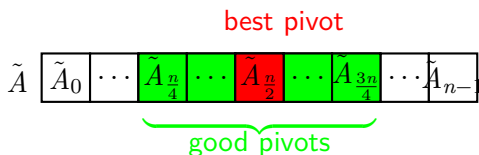
Applying the general DIVIDE AND CONQUER paradigm

SELECT(A, k)

```
1: Choose an element  $A_i$  from  $A$  as a pivot;  
2:  $S_+ = \{\}$ ;  $S_- = \{\}$ ;  
3: for all element  $A_j$  in  $A$  do  
4:   if  $A_j > A_i$  then  
5:      $S_+ = S_+ \cup \{A_j\}$ ;  
6:   else  
7:      $S_- = S_- \cup \{A_j\}$ ;  
8:   end if  
9: end for  
10: if  $|S_-| = k - 1$  then  
11:   return  $A_i$ ;  
12: else if  $|S_-| > k - 1$  then  
13:   return SELECT( $S_-, k$ );  
14: else  
15:   return SELECT( $S_+, k - |S_-| - 1$ );  
16: end if
```

Note: Unlike QUICKSORT, the SELECT algorithm needs to consider only one subproblem. The algorithm would be efficient if the subproblem size, i.e., $\|S_+\|$ or $\|S_-\|$, decreases exponentially as iteration proceeds.

Question: How to choose a pivot?



- Worst choice: select the smallest/largest element as pivot at each iteration. The subproblems decrease **linearly** in size.

$$T(n) = T(n-1) + O(n) = O(n^2)$$

- Best choice: select the **exact median** at each iteration. The subproblems decrease **exponentially** in size.

$$T(n) = T(\frac{n}{2}) + O(n) = O(n)$$

- Good choice: select a **nearly-central element** such that a fixed proportion of elements fall both below and over it, i.e., $\|S_+\| \geq \epsilon n$, and $\|S_-\| \geq \epsilon n$ for a fixed $\epsilon > 0$, say $\epsilon = \frac{1}{4}$. In this case, the subproblems decrease **exponentially** in size, too.

$$\begin{aligned} T(n) &\leq T((1-\epsilon)n) + O(n) \\ &\leq cn + c(1-\epsilon)n + c(1-\epsilon)^2n + \dots \\ &= O(n) \end{aligned}$$

How to efficiently get a **nearly-central** pivot?

- Selection of **nearly-central pivots** always leads to small subproblems, which will speed up the algorithm regardless of k . But how to obtain **nearly-central pivots**?
- We **estimate median of the whole set** through examining a **sample of the whole set**. The following samples have been tried:
 - ① Select a nearly-central pivot via **examining medians of groups**;
 - ② Select a nearly-central pivot via **randomly selecting an element**;
 - ③ Select a nearly-central pivot via **examining a random sample**.
- Note: In 1975, Sedgewick proposed a similar pivot-selecting strategy called **“median-of-three”** for QUICKSORT: selecting the median of the first, middle, and last elements as pivot. The “median-of-three” rule gives a good estimate of the best pivot.

Strategy 1: BFPRT algorithm uses median of medians as pivot

Strategy 1: Median of medians [Blum et al, 1973]

	0	5	6	21	3	17	14	4	1	22	8
	2	9	11	25	16	19	31	20	36	29	18
Medians	7	10	13	26	27	32	34	35	38	42	44
	12	24	23	30	43	33	37	41	46	49	48
	15	51	28	40	45	53	39	47	50	54	52

SELECT(A, k)

- 1: Line up elements in groups of 5 elements;
- 2: Find the median of each group; //Cost $\frac{6}{5}n$ time
- 3: Find the median of medians (denoted as M) through recursively running SELECT over the group medians; // $T(\frac{n}{5})$ time
- 4: Use M as pivot to partition A into S_- and S_+ ; // $O(n)$ time
- 5: **if** $|S_-| = k - 1$ **then**
- 6: **return** M ;
- 7: **else if** $|S_-| > k - 1$ **then**
- 8: **return** SELECT(S_-, k); //at most $T(\frac{7}{10}n)$ time
- 9: **else**
- 10: **return** SELECT($S_+, k - |S_-| - 1$); //at most $T(\frac{7}{10}n)$ time
- 11: **end if**

	0	5	6	21	3	17	14	4	1	22	8
	2	9	11	25	16	19	31	20	36	29	18
Medians	7	10	13	26	27	32	34	35	38	42	44
	12	24	23	30	43	33	37	41	46	49	48
	15	51	28	40	45	53	39	47	50	54	52

- Basic idea: Median of medians $M = 32$ is a perfect approximate median as at least $\frac{3n}{10}$ elements are larger (in red), and at least $\frac{3n}{10}$ elements are smaller than M (in blue). Thus, at least $\frac{3n}{10}$ elements will not appear in S_+ and S_- .
- Running time:

$$T(n) \leq T\left(\frac{n}{5}\right) + T\left(\frac{7n}{10}\right) + O(n) = O(n).$$

Actually it takes at most $24n$ comparisons.

BFPRT algorithm: an in-place implementation

SELECT(A, l, r, k)

```
1: while TRUE do
2:   if  $l == r$  then
3:     return  $l$ ;
4:   end if
5:    $p = \text{PIVOT}(A, l, r)$ ; //Use median of medians  $A[p]$  as pivot ;
6:    $pos = \text{PARTITION}(A, l, r, p)$ ; //  $pos$  represents the final
   position of the pivot,  $A[l..pos - 1]$  deposit  $S_-$  and
    $A[pos + 1..r]$  deposit  $S_+$ ;
7:   if  $(k - 1) == pos$  then
8:     return  $k - 1$ ;
9:   else if  $(k - 1) < pos$  then
10:     $r = pos - 1$ ;
11:  else
12:     $l = pos + 1$ ;
13:  end if
14: end while
```

PIVOT(A, l, r): get median of medians

PIVOT(A, l, r)

```
1: if ( $r - l$ ) < 5 then  
2:   return PARTITION5( $A, l, r$ ); //Get median for 5 or less  
   elements;  
3: end if  
4: for  $i = l$  to  $r$  by 5 do  
5:    $right = i + 4$ ;  
6:   if  $right > r$  then  
7:      $right = r$ ;  
8:   end if  
9:    $m = \text{PARTITION5}(A, i, right)$ ; //Get median of a group;  
10:  Swap  $A[m]$  and  $A[l + \lfloor \frac{i-l}{5} \rfloor]$ ;  
11: end for  
12: return SELECT( $A, l, l + \lfloor \frac{r-l}{5} \rfloor, l + \frac{r-l}{10}$ );
```

PARTITION(A, l, r, p): Partition A into S_- and S_+

PARTITION(A, l, r, p)

- 1: $pivot = A[p]$;
 - 2: Swap $A[p]$ and $A[r]$; //Move pivot to the right end;
 - 3: $i = l$;
 - 4: **for** $j = l$ to $r - 1$ **do**
 - 5: **if** $A[j] < pivot$ **then**
 - 6: Swap $A[i]$ and $A[j]$;
 - 7: $i++$;
 - 8: **end if**
 - 9: **end for**
 - 10: Swap $A[r]$ and $A[i]$;
 - 11: **return** i ;
- Basic idea: Swap $A[p]$ and $A[r]$ to move pivot to the right end first, and then execute the PARTITION function used by Lomuto's QUICKSORT algorithm.

An example: Iteration #1 of SELECT($A, 0, 15, 7$)

8	1	15	10	4	3	2	9	7	12	5	16	14	6	13	11
---	---	----	----	---	---	---	---	---	----	---	----	----	---	----	----

↓ Find group medians

8	1	15	10	4	3	2	9	7	12	5	16	14	6	13	11
---	---	----	----	---	---	---	---	---	----	---	----	----	---	----	----

↓ Swap medians to end

8	7	13	11	4	3	2	9	1	12	5	16	14	6	15	10
---	---	----	----	---	---	---	---	---	----	---	----	----	---	----	----

↓ Find **pivot** using SELECT($A, 0, 3, 2$)

8	7	13	11	4	3	2	9	1	12	5	16	14	6	15	10
---	---	----	----	---	---	---	---	---	----	---	----	----	---	----	----

↓ PARTITION($A, 0, 15, 3$)

8	7	10	4	3	2	9	1	5	6	11	16	14	12	15	13
---	---	----	---	---	---	---	---	---	---	----	----	----	----	----	----

Iteration #2: $\text{SELECT}(A, 0, 9, 7)$

8	7	10	4	3	2	9	1	5	6	11	16	14	12	15	13
---	---	----	---	---	---	---	---	---	---	----	----	----	----	----	----

↓ Find group medians

8	7	10	4	3	2	9	1	5	6	11	16	14	12	15	13
---	---	----	---	---	---	---	---	---	---	----	----	----	----	----	----

↓ Swap medians to end

7	5	10	4	3	2	9	1	8	6	11	16	14	12	15	13
---	---	----	---	---	---	---	---	---	---	----	----	----	----	----	----

↓ Find pivot using $\text{SELECT}(A, 0, 1, 1)$

7	5	10	4	3	2	9	1	8	6	11	16	14	12	15	13
---	---	----	---	---	---	---	---	---	---	----	----	----	----	----	----

↓ $\text{PARTITION}(A, 0, 9, 1)$

4	3	2	1	5	10	9	7	8	6	11	16	14	12	15	13
---	---	---	---	---	----	---	---	---	---	----	----	----	----	----	----

Iteration #3: SELECT($A, 5, 9, 7$)

4	3	2	1	5	10	9	7	8	6	11	16	14	12	15	13
---	---	---	---	---	----	---	---	---	---	----	----	----	----	----	----

↓ Find group medians

4	3	2	1	5	10	9	7	8	6	11	16	14	12	15	13
---	---	---	---	---	----	---	---	---	---	----	----	----	----	----	----

↓ Move medians to end

4	3	2	1	5	8	9	7	10	6	11	16	14	12	15	13
---	---	---	---	---	---	---	---	----	---	----	----	----	----	----	----

↓ Find pivot using SELECT($A, 5, 5, 1$)

4	3	2	1	5	8	9	7	10	6	11	16	14	12	15	13
---	---	---	---	---	---	---	---	----	---	----	----	----	----	----	----

↓ PARTITION($A, 5, 9, 5$)

4	3	2	1	5	6	7	8	10	9	11	16	14	12	15	13
---	---	---	---	---	---	---	---	----	---	----	----	----	----	----	----

Return $A[6] = 7$

Question: How about setting other group size?

- It is easy to prove $T(n) = O(n)$ when setting group size as 7 or larger.
- However, when we setting group size as 3, we have:

$$T(n) \leq T(\frac{n}{3}) + T(\frac{2n}{3}) + O(n) = O(n \log n)$$

- Note that BFPRT algorithm always selects the median of medians as pivot regardless of the value of k . In 2017, Zeng et al. proposed to use fractile of medians rather than median of medians as pivot and selected appropriate fractile of medians according to k .

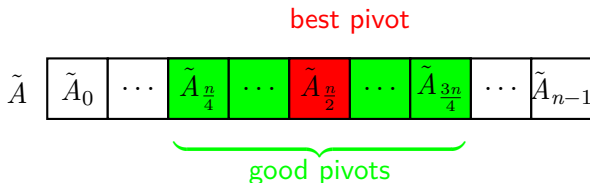
Strategy 2: `QUICKSELECT` algorithm randomly select an element as pivot

Strategy 2: Selecting a pivot randomly [Hoare, 1961]

QUICKSELECT(A, k)

```
1: Choose an element  $A_i$  from  $A$  uniformly at random;  
2:  $S_+ = \{\}$ ;  
3:  $S_- = \{\}$ ;  
4: for all element  $A_j$  in  $A$  do  
5:   if  $A_j > A_i$  then  
6:      $S_+ = S_+ \cup \{A_j\}$ ;  
7:   else  
8:      $S_- = S_- \cup \{A_j\}$ ;  
9:   end if  
10: end for  
11: if  $|S_-| = k - 1$  then  
12:   return  $A_i$ ;  
13: else if  $|S_-| > k - 1$  then  
14:   return QUICKSELECT( $S_-, k$ );  
15: else  
16:   return QUICKSELECT( $S_+, k - |S_-| - 1$ );  
17: end if
```

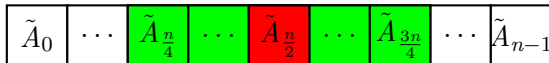
Randomized DIVIDE AND CONQUER cont'd



- Basic idea: when selecting an element uniformly at random, it is highly likely to get a good pivot since a fairly large fraction of the elements are nearly-central.

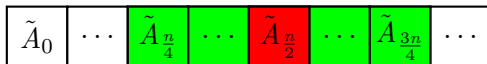
An example

Iteration #1



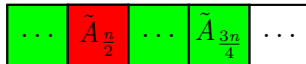
↓ Select \tilde{A}_{n-1} as pivot

Iteration #2



↓ Select $\tilde{A}_{\frac{n}{4}}$ as pivot

Iteration #3



- Selecting a **nearly-central pivot** will lead to a $\frac{3}{4}$ shrinkage of problem size.
- Two iterations are expected before selecting a **nearly-central pivot**.

Theorem

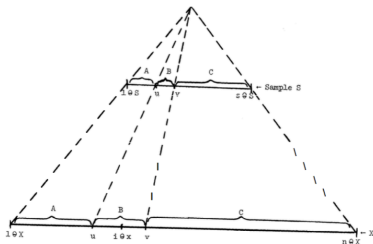
The expected running time of QUICKSELECT is $O(n)$.

Proof.

- We divide the execution into a series of phases: phase j contains a collection of iterations when the size of set under consideration is in $[n(\frac{3}{4})^{j+1} + 1, n(\frac{3}{4})^j]$, say $[\frac{3}{4}n + 1, n]$ for phase 0, and $[\frac{9}{16}n + 1, \frac{3}{4}n]$ for phase 1.
- Let X be the number of comparison that QUICKSELECT uses, and X_j be the number of comparison in phase j . Thus,
$$X = X_0 + X_1 + \dots$$
- Consider phase j . The probability to find a nearly-central pivot is $\frac{1}{2}$ since half elements are nearly-central. Selecting a nearly-central pivot will lead to a $\frac{3}{4}$ shrinkage of problem size and therefore make the execution enter phase $(j+1)$. Thus, the expected iteration number in phase j is 2.
- Each iteration in phase j performs at most $cn(\frac{3}{4})^j$ comparison j since there are at most $n(\frac{3}{4})^j$ elements. Thus, $E[X_j] \leq 2cn(\frac{3}{4})^j$.
- Hence $E[X] = E[X_0 + X_1 + \dots] \leq \sum_j 2cn(\frac{3}{4})^j \leq 8cn$.

Strategy 3: FLOYD-RIVEST algorithm selects a pivot based on random samples

Strategy 3: Selecting pivots according to a random sample



- In 1973, Robert Floyd and Ronald Rivest proposed to select pivot using **random sampling** technique.
- Basic idea: A random sample, if sufficiently large, is a good representation of the whole set. Specifically, the median of a sample is an **unbiased point estimator** of the median of the whole set. We can also use **interval estimation**, i.e., a small interval that is expected to contain the median of the whole set with high probability.

Floyd-Rivest algorithm for SELECTION [1973]

FLOYD-RIVEST-SELECT(A, k)

- 1: Select a small random sample S (with replacement) from A .
 - 2: Select two pivots, denoted as u and v , from S through recursively calling FLOYD-RIVEST-SELECT. The interval $[u, v]$, although small, is expected to cover the k -th smallest element of A .
 - 3: Divide A into three dis-joint subsets: L contains the elements less than u , M contains elements in $[u, v]$, and H contains the elements greater than v .
 - 4: Partition A into these three sets through comparing each element A_i with u and v : if $k \leq \frac{n}{2}$, A_i is compared with v first and then to u only if $A_i \leq v$. The order is reversed if $k > \frac{n}{2}$.
 - 5: The k -th smallest element of A is selected through recursively running over an appropriate subset.
- Here we present a variant of Flyod-Rivest algorithm called LAZYSELECT, which is much easier to analyze.

LAZYSELECTMEDIAN algorithm

LAZYSELECTMEDIAN(A)

- 1: Randomly sample r elements (with replacement) from $A = [A_0, A_1, A_2, \dots, A_{n-1}]$. Denote the sample as S .
- 2: Sort S . Let u be the $\frac{1-\delta}{2}r$ -th smallest element of S and v be the $\frac{1+\delta}{2}r$ -th smallest element of S .
- 3: Divide A into three dis-joint subsets:

$$L = \{A_i : A_i < u\};$$

$$M = \{A_i : u \leq A_i \leq v\};$$

$$H = \{A_i : A_i > v\};$$

- 4: Check the following constraints of M :

- M covers the median: $|L| \leq \frac{n}{2}$ and $|H| \leq \frac{n}{2}$
- M should not be too large: $|M| \leq c\delta n$

If one of the constraints was violated, got to STEP 1.

- 5: Sort M and return the $(\frac{n}{2} - |L|)$ -th smallest of M as the median of A .

An example

Input: A . $n = |A| = 16$. **Set** $\delta = \frac{1}{2}$

8	1	15	10	4	3	2	9	7	12	5	16	14	6	13	11
---	---	----	----	---	---	---	---	---	----	---	----	----	---	----	----

↓ **Sample** $r = 8$ **elements**

8	1	15	10	4	3	2	9	7	12	5	16	14	6	13	11
---	---	----	----	---	---	---	---	---	----	---	----	----	---	----	----

$$S = \{2, 4, 5, 8, 11, 13, 15, 16\}$$

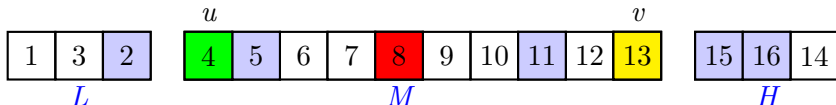
↓ **Divide** A **into** L , M , **and** H

			u											v			
1	3	2	4	5	6	7	8	9	10	11	12	13	15	16	14		
L			M											H			

Return 8 **as the median of** A

Elaborately-designed δ and r

$$S = \{2, 4, 5, 8, 11, 13, 15, 16\}$$



- We expect the following two properties of M :
 - On one side, $|M|$ should be **sufficiently large** such that the median of A is covered by M with high probability.
 - On the other side, $|M|$ should be **sufficiently small** such that the sorting operation in Step 5 will not take a long time.
- We claim that $|M| = \Theta(n^{\frac{3}{4}})$ is an appropriate size that satisfies these two constraints simultaneously.
- To obtain such a M , we set $r = n^{\frac{3}{4}}$, and $\delta = n^{-\frac{1}{4}}$ as M is expected to have a size of $\delta n = n^{\frac{3}{4}}$.

Time-complexity analysis: linear time

LAZYSELECTMEDIAN(A)

- 1: Randomly sample r elements (with replacement) from $A = [A_0, A_1, A_2, \dots, A_{n-1}]$. Denote the sample as S . **//Set $r = n^{\frac{3}{4}}$**
- 2: Sort S . Let u be the $\frac{1-\delta}{2}r$ -th smallest element of S and v be the $\frac{1+\delta}{2}r$ -th smallest element of S . **//Take $O(r \log r) = o(n)$ time**
- 3: Divide A into three dis-joint subsets: **//Take $2n$ steps**

$$L = \{A_i : A_i < u\};$$

$$M = \{A_i : u \leq A_i \leq v\};$$

$$H = \{A_i : A_i > v\};$$

- 4: Check the following constraints of M :

- M covers the median: $|L| \leq \frac{n}{2}$ and $|H| \leq \frac{n}{2}$
- M should not be too large: $|M| \leq c\delta n$

If one of the constraints was violated, got to Step 1.

- 5: Sort M and return the $(\frac{n}{2} - |L|)$ -th smallest of M as the median of A .

//Take $O(\delta n \log(\delta n)) = o(n)$ time when setting $\delta = n^{-\frac{1}{4}}$

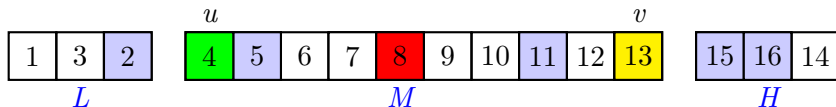
- Total running time (in one pass): $2n + o(n)$. The best known deterministic algorithm takes $3n$ but it is too complicated. On the hand, it has been proved at least $2n$ steps are needed.

Analysis of the success probability in one pass

Theorem

With probability $1 - O(n^{-\frac{1}{4}})$, LAZYSELECTMEDIAN reports the median in the first pass. Thus, the total running time is only $2n + o(n)$.

$$S = \{2, 4, 5, 8, 11, 13, 15, 16\}$$



- There are two types of failures in one pass, namely, M does not cover the median of the whole set A , and M is too large. We claim that the probability of both types of failures are as small as $O(n^{-\frac{1}{4}})$. Here we present proof for the first type only.

M covers the median of A with high probability

- We argue that $|L| > \frac{n}{2}$ occurs with probability $O(n^{-\frac{1}{4}})$. Note that $|L| > \frac{n}{2}$ implies that u is greater than the median of A , and thus at least $\frac{1+\delta}{2}r$ elements in S are greater than the median.
- Let $X = x_1 + x_2 + \dots + x_r$ be the number of sampled elements greater than the median of A , where x_i is an index variable:
$$x_i = \begin{cases} 1 & \text{if the } i\text{-th element in } S \text{ is greater than the median} \\ 0 & \text{otherwise} \end{cases}$$
- Then $E(x_i) = \frac{1}{2}$, $\sigma^2(x_i) = \frac{1}{4}$, $E(X) = \frac{1}{2}r$, $\sigma^2(X) = \frac{1}{4}r$, and

$$\Pr(|L| > \frac{n}{2}) \leq \Pr(X \geq \frac{1+\delta}{2}r) \quad (2)$$

$$= \frac{1}{2} \Pr(|X - E(X)| \geq \frac{\delta}{2}r) \quad (3)$$

$$\leq \frac{\frac{1}{2} \sigma^2(X)}{(\frac{\delta}{2}r)^2} \quad (4)$$

$$= \frac{1}{2} \frac{1}{\delta^2 r} \quad (5)$$

$$= \frac{1}{2} n^{-\frac{1}{4}} \quad (6)$$

MULTIPLICATION problem: to multiply **two n -bits integers**

MULTIPLICATION problem

INPUT: Two n -bits integers x and y . Here we represent x as an array $x_0x_1\dots x_{n-1}$, where x_i denotes the i -th bit of x . Similarly, we represent y as an array $y_0y_1\dots y_{n-1}$, where y_i denotes the i -th bit of y .

OUTPUT: The product $x \times y$.

- An example:

$$\begin{array}{r} 12 \\ \times 34 \\ \hline 48 \\ 36 \\ \hline 408 \end{array}$$

- Question: Is the grade-school $O(n^2)$ algorithm optimal?



- Conjecture: In 1960, Andrey Kolmogorov conjectured that any algorithm for that task would require $\Omega(n^2)$ elementary operations.

MULTIPLICATION problem: Trial 1

- Key observation: both x and y can be decomposed into two parts;
- DIVIDE AND CONQUER:
 - 1 **Divide:** $x = x_h \times 2^{\frac{n}{2}} + x_l$, $y = y_h \times 2^{\frac{n}{2}} + y_l$,
 - 2 **Conquer:** calculate $x_h y_h$, $x_h y_l$, $x_l y_h$, and $x_l y_l$;
 - 3 **Combine:**

$$xy = (x_h \times 2^{\frac{n}{2}} + x_l)(y_h \times 2^{\frac{n}{2}} + y_l) \quad (7)$$

$$= x_h y_h 2^n + (x_h y_l + x_l y_h) 2^{\frac{n}{2}} + x_l y_l \quad (8)$$

MULTIPLICATION problem: Trial 1

- Example:
 - Objective: to calculate 12×34
 - $x = 12 = 1 \times 10 + 2$, $y = 34 = 3 \times 10 + 4$
 - $x \times y = (1 \times 3) \times 10^2 + ((1 \times 4) + (2 \times 3)) \times 10 + 2 \times 4$
- Note: 4 sub-problems, 3 additions, and 2 shifts;
- Time-complexity: $T(n) = 4T(\frac{n}{2}) + O(n) = O(n^2)$

Question: can we reduce the number of sub-problems?

Reduce the number of sub-problems

\times	y_h	y_l
x_h	$x_h y_h$	$x_h y_l$
x_l	$x_l y_h$	$x_l y_l$

- Our objective is to calculate $x_h y_h 2^n + (x_h y_l + x_l y_h) 2^{\frac{n}{2}} + x_l y_l$.
- Thus it is unnecessary to calculate $x_h y_l$ and $x_l y_h$ separately; we just need to calculate the sum $(x_h y_l + x_l y_h)$.
- It is obvious that
$$(x_h y_l + x_l y_h) + x_h y_h + x_l y_l = (x_h + x_l) \times (y_h + y_l).$$
- The sum $(x_h y_l + x_l y_h)$ can be calculated using only **one** additional multiplication.
- This idea is dated back to Carl. F. Gauss: Calculation of the product of two complex numbers
$$(a + bi)(c + di) = (ac - bd) + (bc + ad)i$$
 seems to require four multiplications, three multiplications ac , bd , and $(a + b)(c + d)$ are sufficient because $bc + ad = (a + b)(c + d) - ac - bd$.

MULTIPLICATION problem: a clever **conquer**

[Karatsuba-Ofman, 1962]



Figure 3: Anatolii Alexeevich Karatsuba

- Karatsuba algorithm was the first multiplication algorithm asymptotically faster than the quadratic "grade school" algorithm.

MULTIPLICATION problem: a clever conquer

- DIVIDE AND CONQUER:

- ➊ **Divide:** $x = x_h \times 2^{\frac{n}{2}} + x_l$, $y = y_h \times 2^{\frac{n}{2}} + y_l$,
- ➋ **Conquer:** calculate $x_h y_h$, $x_l y_l$, and $P = (x_h + x_l)(y_h + y_l)$;
- ➌ **Combine:**

$$xy = (x_h \times 2^{\frac{n}{2}} + x_l)(y_h \times 2^{\frac{n}{2}} + y_l) \quad (9)$$

$$= x_h y_h 2^n + (x_h y_l + x_l y_h) 2^{\frac{n}{2}} + x_l y_l \quad (10)$$

$$= x_h y_h 2^n + (P - x_h y_h - x_l y_l) 2^{\frac{n}{2}} + x_l y_l \quad (11)$$

Karatsuba-Ofman algorithm

- Example:
 - Objective: to calculate 12×34
 - $x = 12 = 1 \times 10 + 2$, $y = 34 = 3 \times 10 + 4$
 - $P = (1 + 2) \times (3 + 4)$
 - $x \times y = (1 \times 3) \times 10^2 + (P - 1 \times 3 - 2 \times 4) \times 10 + 2 \times 4$
- Note: 3 sub-problems, 6 additions, and 2 shifts;
- Time-complexity:
$$T(n) = 3T\left(\frac{n}{2}\right) + cn = O(n^{\log_2 3}) = O(n^{1.585})$$
- Karatsuba algorithm is a special case of Toom-Cook algorithm. Toom-3 algorithm decomposes both x and y into 3 parts, and calculates xy in $O(n^{1.465})$ time.

Theoretical analysis vs. empirical performance

- For large n , Karatsuba's algorithm will perform fewer shifts and single-digit additions.
- For small values of n , however, the extra shift and add operations may make it run slower.
- The crossover point depends on the computer platform and context.
- When applying FFT technique over ring, the MULTIPLICATION can be finished in $O(n \log n \log \log n)$ time.

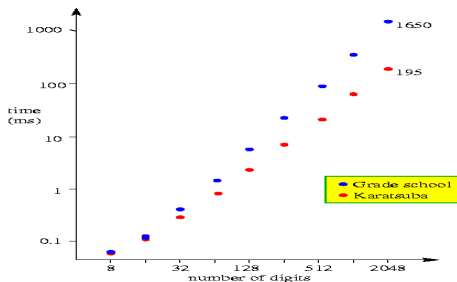


Figure 4: Sun SPARC4, g++ -O4, random input. See

- Problem: Given two n -digit numbers s and t , to calculate $q = s/t$ and $r = s \bmod t$.
- Method:
 - 1 Calculate $x = 1/t$ using Newton's method first:
$$x_{i+1} = 2x_i - t \times x_i^2$$
 - 2 At most $\log n$ iterations are needed.
 - 3 Thus division is as fast as multiplication.

Details of FAST DIVISION: Newton's method

- Objective: Calculate $x = 1/t$.
 - x is the root of $f(x) = 0$, where $f(x) = (t - \frac{1}{x})$. (Why the form here?)
 - Newton's method:

$$x_{i+1} = x_i - \frac{f(x_i)}{f'(x_i)} \quad (12)$$

$$= x_i - \frac{t - \frac{1}{x_i}}{\frac{1}{x_i^2}} \quad (13)$$

$$= -t \times x_i^2 + 2x_i \quad (14)$$

- Convergence speed: quadratic, i.e. $\epsilon_{i+1} \leq M\epsilon_i^2$, where M is a supremum of a ratio, and ϵ_i denotes the distance between x_i and $\frac{1}{t}$. Thus the number of iterations is limited by $\log \log t = O(\log n)$.

FAST DIVISION: an example

- Objective: to calculate $\frac{1}{13}$.

#Iteration	x_i	ϵ_i
0	0.018700	-0.058223
1	0.032854	-0.044069
2	0.051676	-0.025247
3	0.068636	-0.008286
4	0.076030	-0.000892
5	0.076912	-1.03583e-05
6	0.076923	-1.39483e-09
7	0.076923	-2.77556e-17
8

- Note: the quadratic convergence implies that the error ϵ_i has a form of $O(e^{2^i})$; thus the iteration number is limited by $\log \log(t)$.

MATRIX MULTIPLICATION problem: to multiply two **matrices**

MATRIX MULTIPLICATION problem

INPUT: Two $n \times n$ matrices A and B ,

$$A = \begin{bmatrix} a_{11} & a_{12} & \cdots & a_{1n} \\ a_{21} & a_{22} & \cdots & a_{2n} \\ \cdots & \cdots & \cdots & \cdots \\ a_{n1} & a_{n2} & \cdots & a_{nn} \end{bmatrix}, B = \begin{bmatrix} b_{11} & b_{12} & \cdots & b_{1n} \\ b_{21} & b_{22} & \cdots & b_{2n} \\ \cdots & \cdots & \cdots & \cdots \\ b_{n1} & b_{n2} & \cdots & b_{nn} \end{bmatrix}$$

OUTPUT: The product $C = AB$.

Grade-school algorithm: $O(n^3)$.

MATRIX MULTIPLICATION problem: Trial 1 I

- Matrix multiplication: Given two $n \times n$ matrices A and B , compute $C = AB$;
 - Grade-school: $O(n^3)$.
- Key observation: matrix can be decomposed into four $\frac{n}{2} \times \frac{n}{2}$ matrices;
- DIVIDE AND CONQUER:
 - Divide:** divide A , B , and C into sub-matrices;
 - Conquer:** calculate products of sub-matrices;
 - Combine:**

$$\begin{bmatrix} C_{11} & C_{12} \\ C_{21} & C_{22} \end{bmatrix} = \begin{bmatrix} A_{11} & A_{12} \\ A_{21} & A_{22} \end{bmatrix} \begin{bmatrix} B_{11} & B_{12} \\ B_{21} & B_{22} \end{bmatrix}$$

$$C_{11} = (A_{11} \times B_{11}) + (A_{12} \times B_{21})$$

$$C_{12} = (A_{11} \times B_{12}) + (A_{12} \times B_{22})$$

$$C_{21} = (A_{21} \times B_{11}) + (A_{22} \times B_{21})$$

$$C_{22} = (A_{21} \times B_{12}) + (A_{22} \times B_{22})$$

- We need to solve 8 sub-problems, and 4 additions; each addition takes $O(n^2)$ time.
- $T(n) = 8T(\frac{n}{2}) + cn^2 = O(n^3)$

Question: can we reduce the number of sub-problems?



Figure 5: Volker Strassen, 2009

- The first algorithm for performing matrix multiplication faster than the $O(n^3)$ time bound.

MATRIX MULTIPLICATION problem: a clever conquer I

- Matrix multiplication: Given two $n \times n$ matrices A and B , compute $C = AB$;
 - Grade-school: $O(n^3)$.
 - Key observation: matrix can be decomposed into four $\frac{n}{2} \times \frac{n}{2}$ matrices;

DIVIDE AND CONQUER:

- Divide:** divide A , B , and C into sub-matrices;
- Conquer:** calculate products of sub-matrices;
- Combine:**

$$\begin{bmatrix} C_{11} & C_{12} \\ C_{21} & C_{22} \end{bmatrix} = \begin{bmatrix} A_{11} & A_{12} \\ A_{21} & A_{22} \end{bmatrix} \begin{bmatrix} B_{11} & B_{12} \\ B_{21} & B_{22} \end{bmatrix}$$

MATRIXMULTIPLICATION problem: a clever conquer II

$$P_1 = A_{11} \times (B_{12} - B_{22}) \quad (15)$$

$$P_2 = (A_{11} + A_{12}) \times B_{22} \quad (16)$$

$$P_3 = (A_{21} + A_{22}) \times B_{11} \quad (17)$$

$$P_4 = A_{22} \times (B_{21} - B_{11}) \quad (18)$$

$$P_5 = (A_{11} + A_{22}) \times (B_{11} + B_{22}) \quad (19)$$

$$P_6 = (A_{12} - A_{22}) \times (B_{21} + B_{22}) \quad (20)$$

$$P_7 = (A_{11} - A_{21}) \times (B_{11} + B_{12}) \quad (21)$$

$$C_{11} = P_4 + P_5 + P_6 - P_2 \quad (22)$$

$$C_{12} = P_1 + P_2 \quad (23)$$

$$C_{21} = P_3 + P_4 \quad (24)$$

$$C_{22} = P_1 + P_5 - P_3 - P_7 \quad (25)$$

- We need to solve 7 sub-problems, and 18 additions/subtraction; each addition/subtraction takes $O(n^2)$ time.
- $T(n) = 7T(\frac{n}{2}) + cn^2 = O(n^{\log_2 7}) = O(n^{2.807})$

- For large n , Strassen algorithm is faster than grade-school method.¹
- Strassen algorithm can be used to solve other problems, say matrix inversion, determinant calculation, finding triangles in graphs, etc.
- Gaussian elimination is not optimal.

¹This heavily depends on the system, including memory access property, hardware design, etc.

- Strassen algorithm performs better than grade-school method only for large n .
- The reduction in the number of arithmetic operations however comes at the price of a somewhat reduced numerical stability,
- The algorithm also requires significantly more memory compared to the naive algorithm.

Fast matrix multiplication

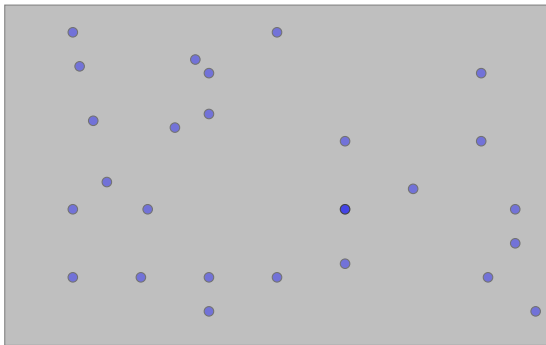
- multiply two 2×2 matrices: 7 scalar sub-problems:
 $O(n^{\log_2 7}) = O(n^{2.807})$ [Strassen 1969]
- multiply two 2×2 matrices: 6 scalar sub-problems:
 $O(n^{\log_2 6}) = O(n^{2.585})$ (impossible)[Hopcroft and Kerr 1971]
- multiply two 3×3 matrices: 21 scalar sub-problems:
 $O(n^{\log_3 21}) = O(n^{2.771})$ (impossible)
- multiply two 20×20 matrices: 4460 scalar sub-problems:
 $O(n^{\log_{20} 4460}) = O(n^{2.805})$
- multiply two 48×48 matrices: 47217 scalar sub-problems:
 $O(n^{\log_{48} 47217}) = O(n^{2.780})$
- Best known till 2010: $O(n^{2.376})$ [Coppersmith-Winograd, 1987]
- Conjecture: $O(n^{2+\epsilon})$ for any $\epsilon > 0$

CLOSESTPAIR problem: given a **set** of points in a plane, to find the closest pair

CLOSESTPAIR problem

INPUT: n points in a plane;

OUTPUT: The pair with the least Euclidean distance.



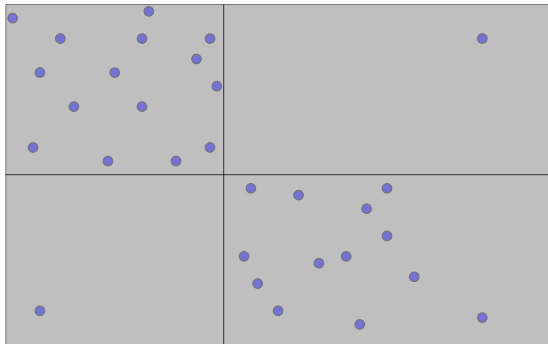
About CLOSESTPAIR problem

- Computational geometry: M. Shamos and D. Hoey were working out efficient algorithm for basic computational primitive in CG in 1970's. They asked a question: does there exist an algorithm using less than $O(n^2)$ time?
- 1D case: it is easy to solve the problem in $O(n \log n)$ via sorting.
- 2D case: a brute-force algorithm works in $O(n^2)$ time by checking all possible pairs.
- **Question:** can we find a faster method?

Trial 1: Divide into 4 subsets

Trial 1: DIVIDE AND CONQUER (4 subsets)

- DIVIDE AND CONQUER: divide into 4 subsets.

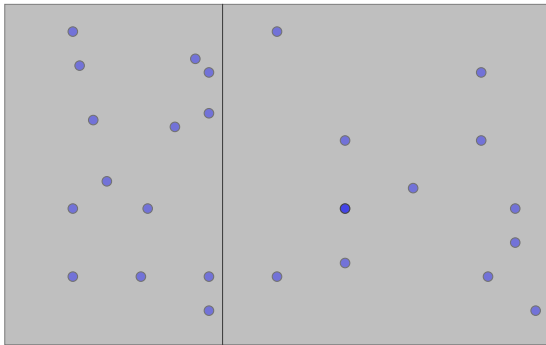


- Difficulties:
 - The subsets might be unbalanced — we cannot guarantee that each subset has approximately $\frac{n}{4}$ points.
 - Since the closest pair might lie in different subsets, we need to consider all $\binom{4}{2}$ pairs of subsets to avoid missing the closest pair, thus complicating the “combine” step.

Trial 2: Divide into 2 halves

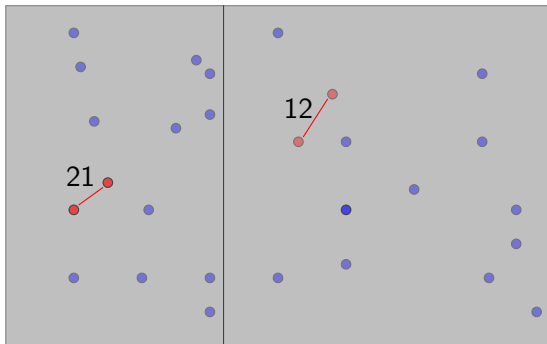
Trial 2: DIVIDE AND CONQUER (2 subsets)

- **Divide:** divide into two halves with equal size.
It is easy to achieve this through sorting by x coordinate first, and then select the median as pivot.



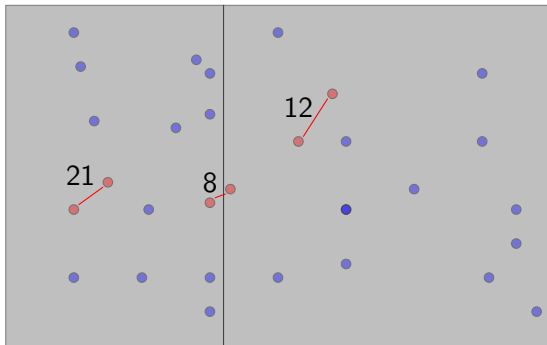
Trial 2: DIVIDE AND CONQUER (2 subsets)

- **Divide:** dividing into two (roughly equal) subsets;
- **Conquer:** finding closest pairs in each half;

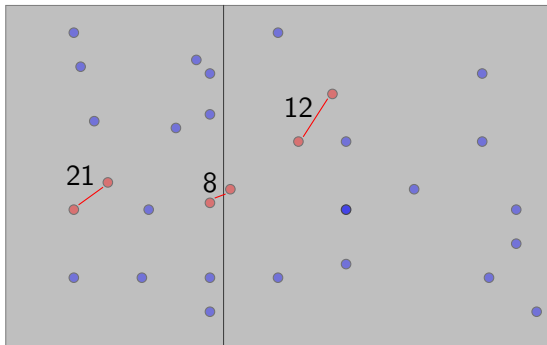


Trial 2: DIVIDE AND CONQUER (2 subsets)

- **Combine:** It suffices to consider the pairs consisting of one point from left half and one point from right half. Simply examining all such pairs will take $O(n^2)$ time.



Two types of redundancy

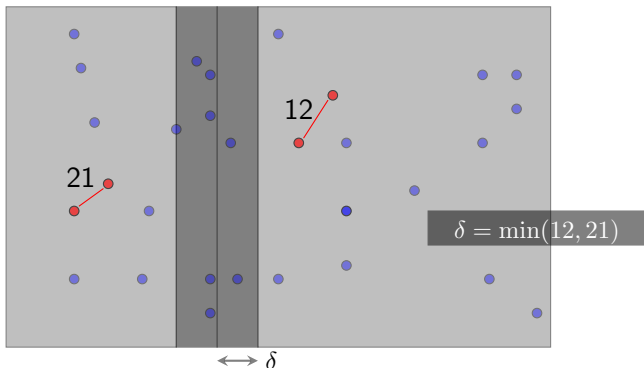


- It is redundant to calculate distance between p_i and p_j if
 - $|x_i - x_j| \geq 12$, or
 - $|y_i - y_j| \geq 12$

Remove redundancy of type 1

- **Observation 1:**

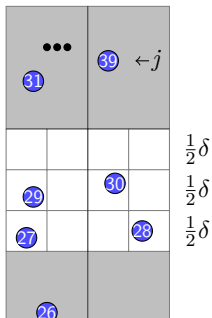
- The third type occurs in **a narrow strip** only; thus, it suffices to check point pairs within the 2δ -strip.
- Here, δ is the minimum of $\text{CLOSESTPAIR}(\text{LEFTHALF})$ and $\text{CLOSESTPAIR}(\text{RIGHTHALF})$.



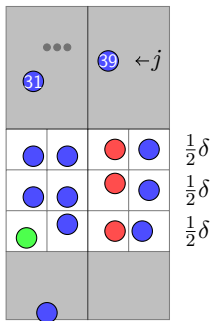
Remove redundancy of type 2

- **Observation 2:**

- Moreover, it is unnecessary to explore **all** point pairs within the 2δ -strip. In fact, for each point p_i , it suffices to examine 11 points for possible closest partners.
- Let's divide the 2δ -strip into grids (size: $\frac{\delta}{2} \times \frac{\delta}{2}$). A grid contains **at most one** point.
- If two points are 2 rows apart, the distance between them should be over δ and thus cannot form closest pair.
- Example: For point 27, it suffices to search within 2 rows for possible closest partners ($< \delta$).

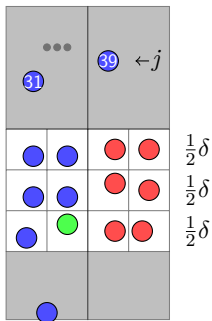


To detect potential closest pair: Case 1



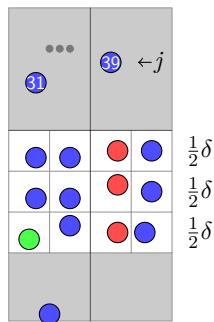
- Green: point i ;
- Red: the possible closest partner (distance $< \delta$) of point i ;

To detect potential closest pair: Case 2



- Green: point i ;
- Red: the possible closest partner (distance $< \delta$) of point i ;

To detect potential closest pair



- If all points within the strip were sorted by y -coordinates, it suffices to calculate distance between each point with its next 11 neighbors.
- Why 11 points here? All red points fall into the subsequent 11 points.

CLOSESTPAIR algorithm

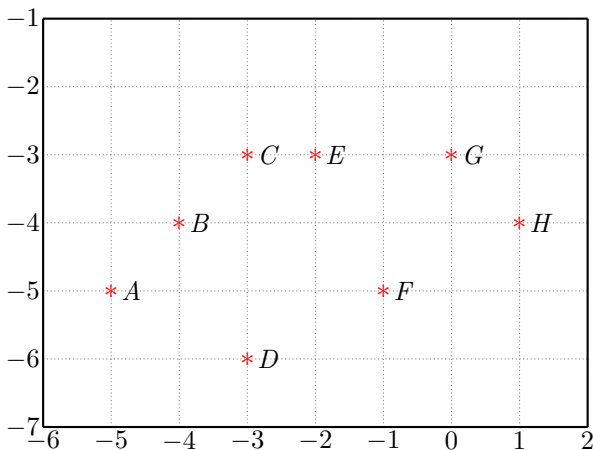
CLOSESTPAIR(p_l, \dots, p_r)

- 1: //To find the closest points within (p_l, \dots, p_r) . Here we assume that p_l, \dots, p_r have already been sorted according to x -coordinate;
 - 2: **if** $r - l == 1$ **then**
 - 3: **return** $d(p_l, p_r)$;
 - 4: **end if**
 - 5: Use the x -coordinate of $p_{\lfloor \frac{l+r}{2} \rfloor}$ to divide p_l, \dots, p_r into two halves;
 - 6: $\delta_1 = \text{CLOSESTPAIR}(\text{LEFTHALF})$; // $T(\frac{n}{2})$
 - 7: $\delta_2 = \text{CLOSESTPAIR}(\text{RIGHTHALF})$; // $T(\frac{n}{2})$
 - 8: $\delta = \min(\delta_1, \delta_2)$;
 - 9: Sort points within the 2δ wide strip by y -coordinate; // $O(n \log n)$
 - 10: Scan points in y -order and calculate distance between each point with its next 11 neighbors. Update δ if finding a distance less than δ ;
 // $O(n)$
- Find closest pair within p_0, p_1, \dots, p_{n-1} :
 CLOSESTPAIR(p_0, \dots, p_{n-1})
 - Time-complexity: $T(n) = 2T(\frac{n}{2}) + O(n \log n) = O(n \log^2 n)$.

CLOSESTPAIR algorithm: improvement

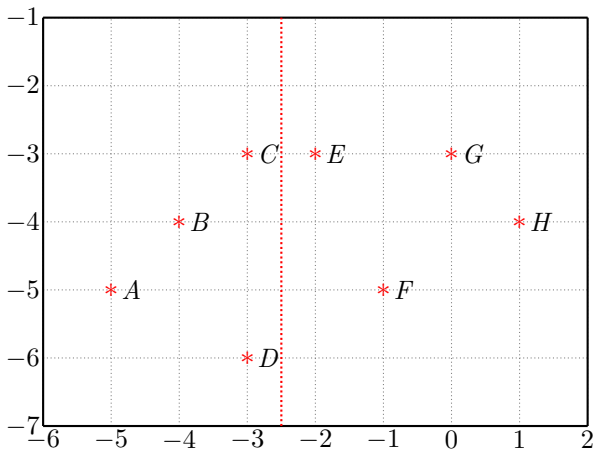
- Note that if the points within the 2δ -wide strip have no structure, we have to sort them from the scratch, which will take $O(n \log n)$ time.
- Let's try to introduce some structure into the points within the 2δ -wide: If the point within each δ -wide strip were already sorted, it is relatively easy to sort the points within the 2δ -wide strip. Specifically,
 - Each recursion keeps two sorted list: one list by x , and the other list by y .
 - We merge two pre-sorted lists into a list as MERGESORT does, which costs only $O(n)$ time.
- Time-complexity: $T(n) = 2T(\frac{n}{2}) + O(n) = O(n \log n)$.

CLOSESTPAIR: an example with 8 points



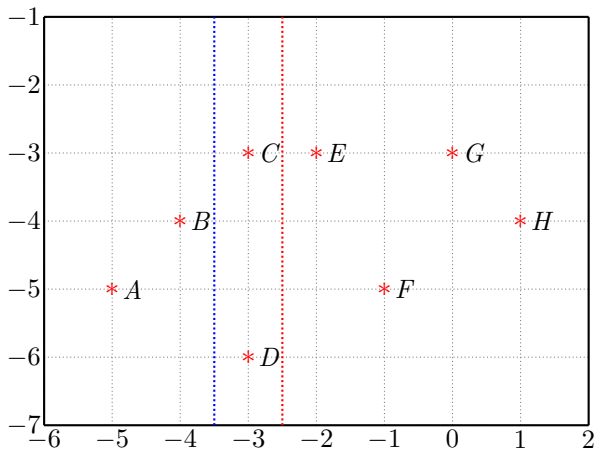
- Objective: to find the closest pair among these 8 points.

CLOSESTPAIR: an example with 8 points

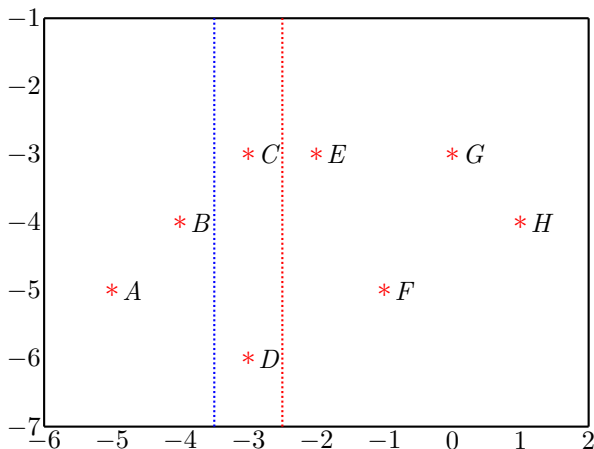


- Objective: to find the closest pair among these 8 points.

Left half: A, B, C, D

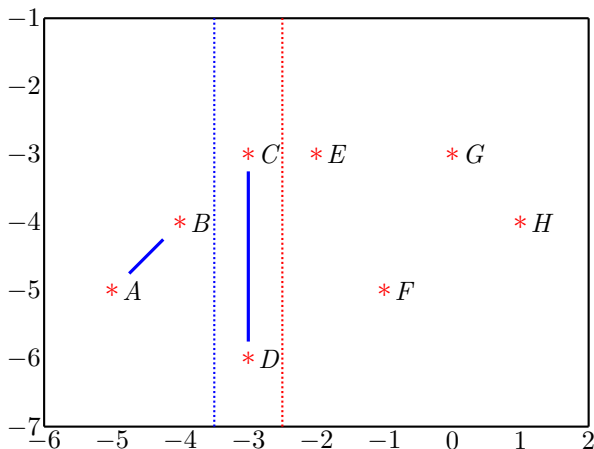


Left half: A, B, C, D



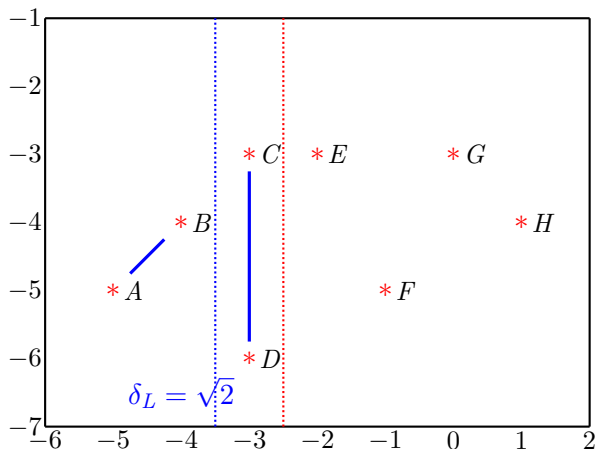
- Pair 1: $d(A, B) = \sqrt{2}$;
- Pair 2: $d(C, D) = 3$; $\Rightarrow \min = \sqrt{2}$; Thus, it suffices to calculate:
- Pair 3: $d(B, C) = \sqrt{2}$;
- Pair 4: $d(B, D) = \sqrt{5}$; $\Rightarrow \delta_L = \sqrt{2}$.

Left half: A, B, C, D



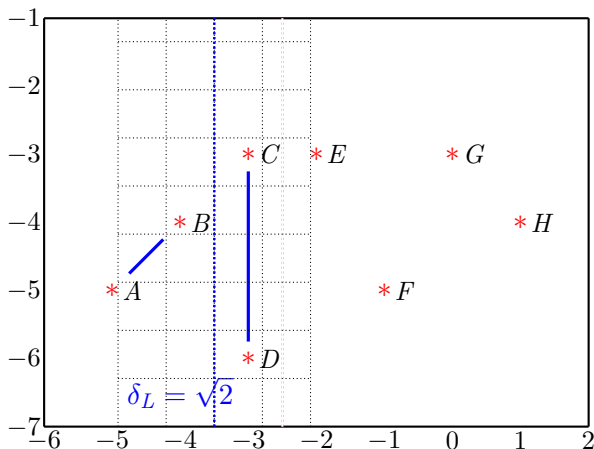
- Pair 1: $d(A, B) = \sqrt{2}$;
- Pair 2: $d(C, D) = 3$; $\Rightarrow \min = \sqrt{2}$; Thus, it suffices to calculate:
- Pair 3: $d(B, C) = \sqrt{2}$;
- Pair 4: $d(B, D) = \sqrt{5}$; $\Rightarrow \delta_L = \sqrt{2}$.

Left half: A, B, C, D



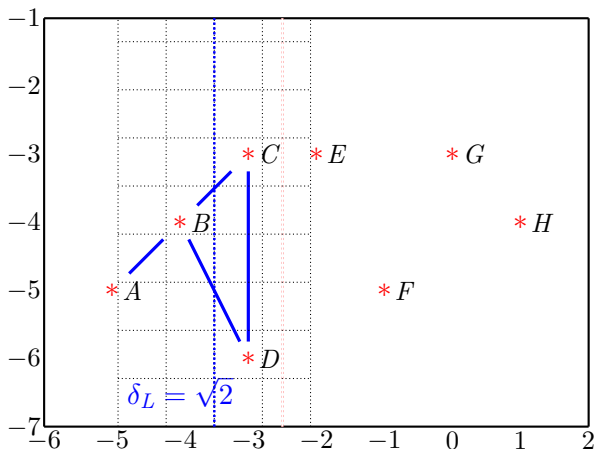
- Pair 1: $d(A, B) = \sqrt{2}$;
- Pair 2: $d(C, D) = 3$; $\Rightarrow \min = \sqrt{2}$; Thus, it suffices to calculate:
- Pair 3: $d(B, C) = \sqrt{2}$;
- Pair 4: $d(B, D) = \sqrt{5}$; $\Rightarrow \delta_L = \sqrt{2}$.

Left half: A, B, C, D



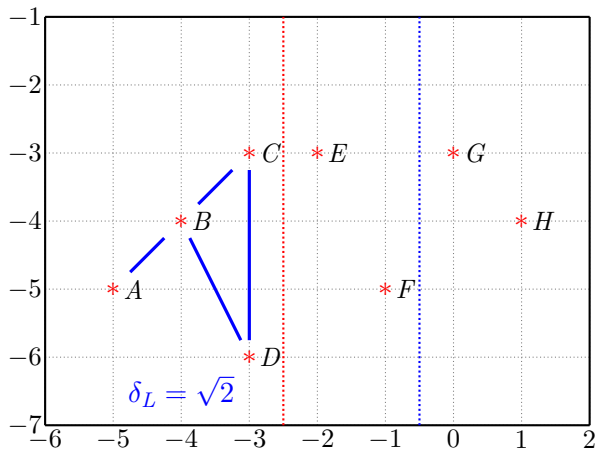
- Pair 1: $d(A, B) = \sqrt{2}$;
- Pair 2: $d(C, D) = 3$; $\Rightarrow \min = \sqrt{2}$; Thus, it suffices to calculate:
- Pair 3: $d(B, C) = \sqrt{2}$;
- Pair 4: $d(B, D) = \sqrt{5}$; $\Rightarrow \delta_L = \sqrt{2}$.

Left half: A, B, C, D

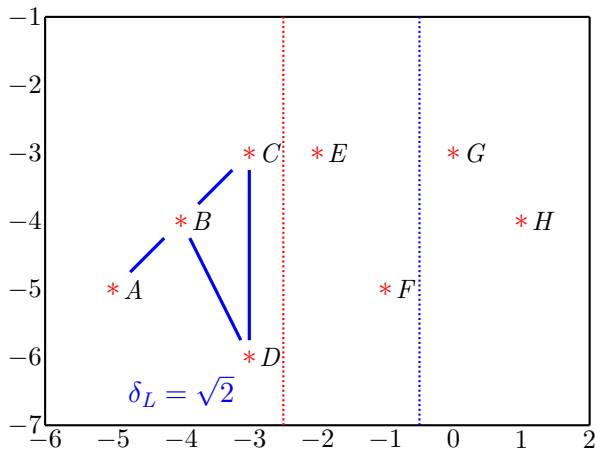


- Pair 1: $d(A, B) = \sqrt{2}$;
- Pair 2: $d(C, D) = 3$; $\Rightarrow \min = \sqrt{2}$; Thus, it suffices to calculate:
- Pair 3: $d(B, C) = \sqrt{2}$;
- Pair 4: $d(B, D) = \sqrt{5}$; $\Rightarrow \delta_L = \sqrt{2}$.

Right half: E, F, G, H

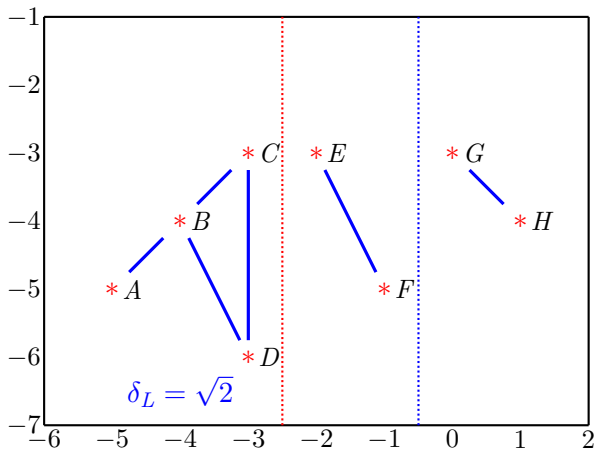


Right half: E, F, G, H



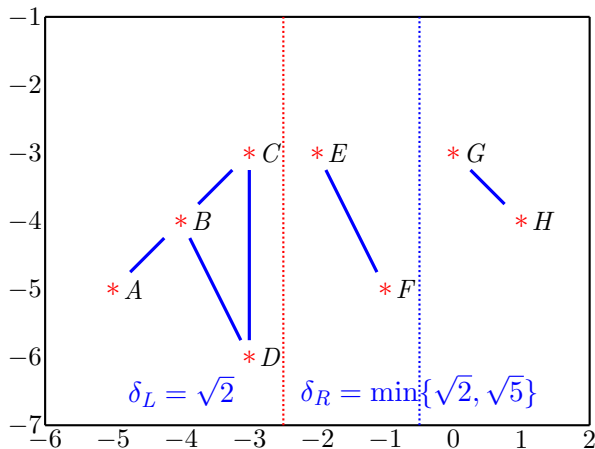
- Pair 5: $d(E, F) = \sqrt{5}$;
- Pair 6: $d(G, H) = \sqrt{2}$; $\Rightarrow \min = \sqrt{2}$; Thus, it suffices to calculate:
- Pair 7: $d(G, F) = \sqrt{5}$; $\Rightarrow \delta_R = \sqrt{2}$.

Right half: E, F, G, H



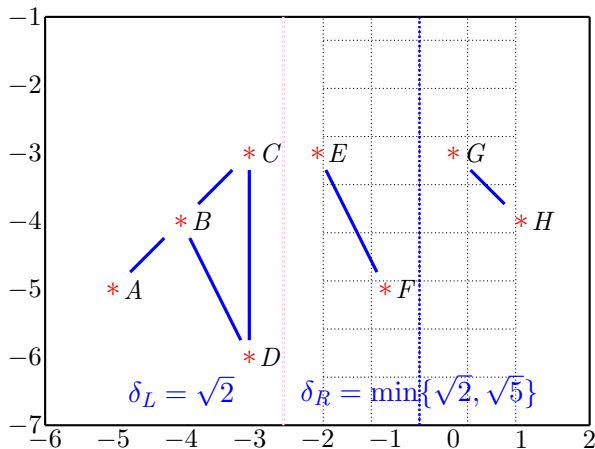
- Pair 5: $d(E, F) = \sqrt{5}$;
- Pair 6: $d(G, H) = \sqrt{2}$; $\Rightarrow \min = \sqrt{2}$; Thus, it suffices to calculate:
- Pair 7: $d(G, F) = \sqrt{5}$; $\Rightarrow \delta_R = \sqrt{2}$.

Right half: E, F, G, H



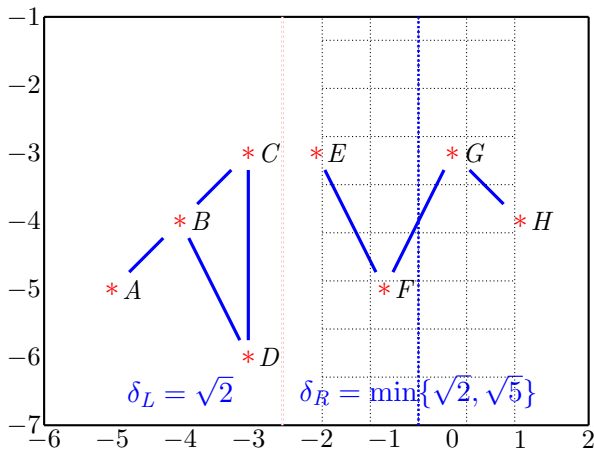
- Pair 5: $d(E, F) = \sqrt{5}$;
- Pair 6: $d(G, H) = \sqrt{2}$; $\Rightarrow \min = \sqrt{2}$; Thus, it suffices to calculate:
- Pair 7: $d(G, F) = \sqrt{5}$; $\Rightarrow \delta_R = \sqrt{2}$.

Right half: E, F, G, H



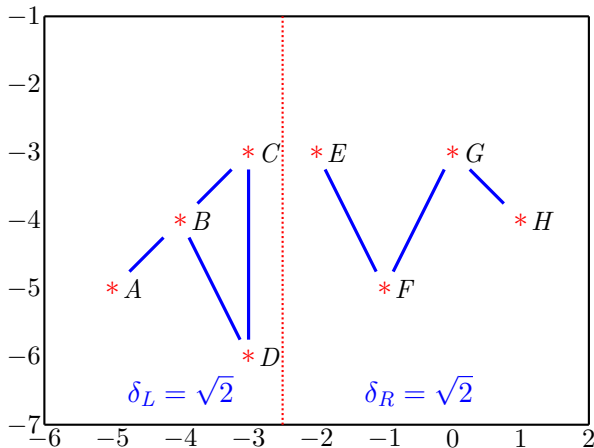
- Pair 5: $d(E, F) = \sqrt{5}$;
- Pair 6: $d(G, H) = \sqrt{2}$; $\Rightarrow \min = \sqrt{2}$; Thus, it suffices to calculate:
- Pair 7: $d(G, F) = \sqrt{5}$; $\Rightarrow \delta_R = \sqrt{2}$.

Right half: E, F, G, H



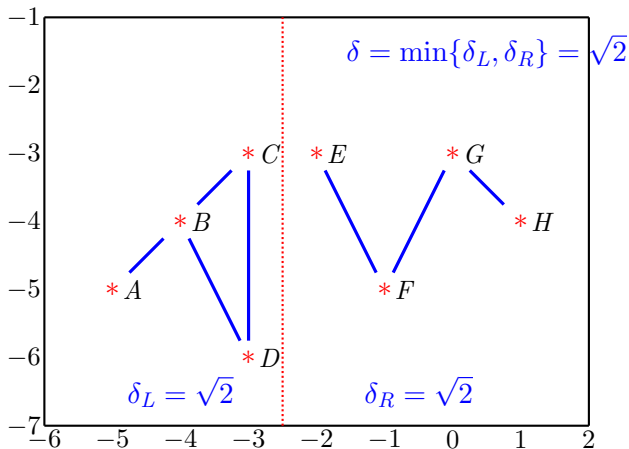
- Pair 5: $d(E, F) = \sqrt{5}$;
- Pair 6: $d(G, H) = \sqrt{2}$; $\Rightarrow \min = \sqrt{2}$; Thus, it suffices to calculate:
- Pair 7: $d(G, F) = \sqrt{5}$; $\Rightarrow \delta_R = \sqrt{2}$.

The entire set: A, B, C, D, E, F, G, H



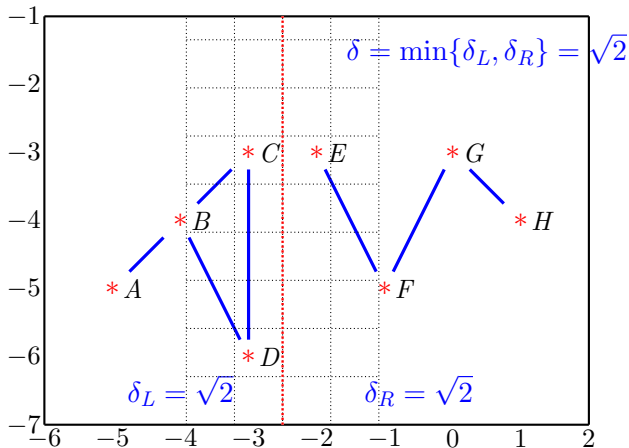
- Pair 8: $d(C, E) = 1$;
- Pair 9: $d(D, E) = \sqrt{10}$; $\Rightarrow \delta = 1$.

The entire set: A, B, C, D, E, F, G, H



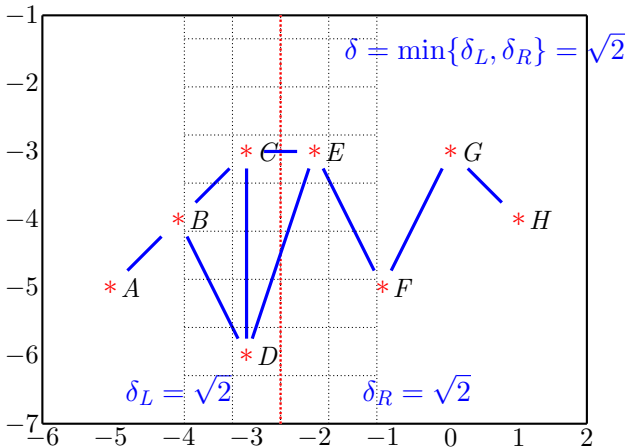
- Pair 8: $d(C, E) = 1$;
- Pair 9: $d(D, E) = \sqrt{10}$; $\Rightarrow \delta = 1$.

The entire set: A, B, C, D, E, F, G, H



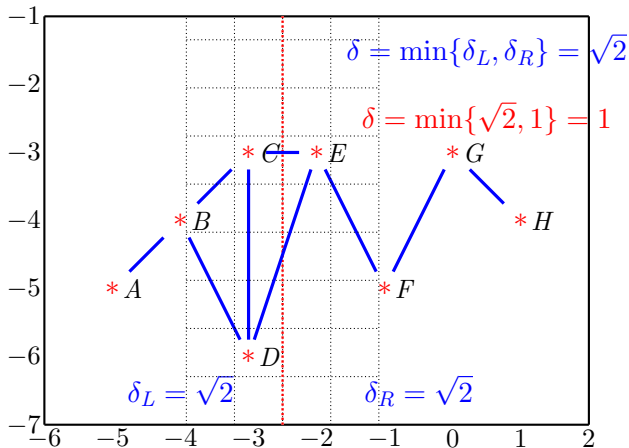
- Pair 8: $d(C, E) = 1$;
- Pair 9: $d(D, E) = \sqrt{10}$; $\Rightarrow \delta = 1$.

The entire set: A, B, C, D, E, F, G, H



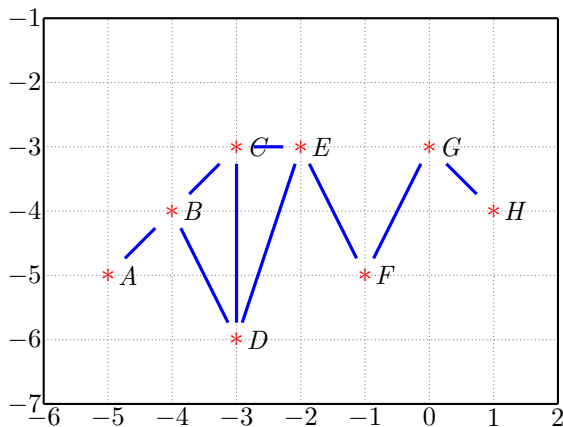
- Pair 8: $d(C, E) = 1$;
- Pair 9: $d(D, E) = \sqrt{10}$; $\Rightarrow \delta = 1$.

The entire set: A, B, C, D, E, F, G, H



- Pair 8: $d(C, E) = 1$;
- Pair 9: $d(D, E) = \sqrt{10}$; $\Rightarrow \delta = 1$.

From $O(n^2)$ to $O(n \log n)$, what did we save?

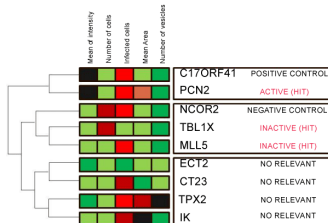


- We calculated distances for only 9 pairs of points (see 'blue' line). The other 19 pairs are redundant due to:
 - at least one of the two points lies out of 2δ -strip.
 - although two points appear in the same 2δ -strip, they are at least 2 rows of grids (size: $\frac{\delta}{2} \times \frac{\delta}{2}$) apart.

Extension: arbitrary (not necessarily geometric) distance functions

Theorem

We can perform bottom-up hierarchical clustering, for any cluster distance function computable in constant time from the distances between subclusters, in total time $O(n^2)$. We can perform median, centroid, Ward, or other bottom-up clustering methods in which clusters are represented by objects, in time $O(n^2 \log^2 n)$ and space $O(n)$.



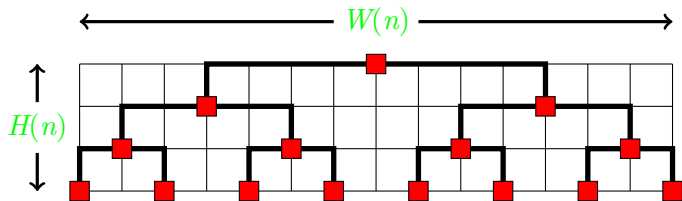
(See Eppstein 1998 for details.)

VLSI embedding: to embed **a tree**

Embedding a tree

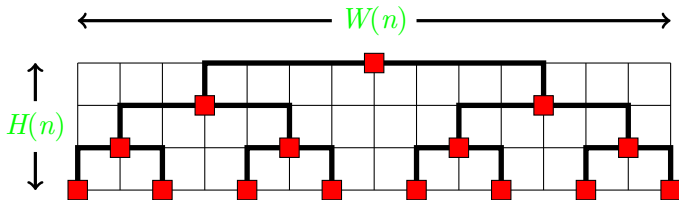
INPUT: Given a binary tree with n node;

OUTPUT: Embedding the tree into a VLSI with minimum area.



Trial 1: divide into two sub-trees

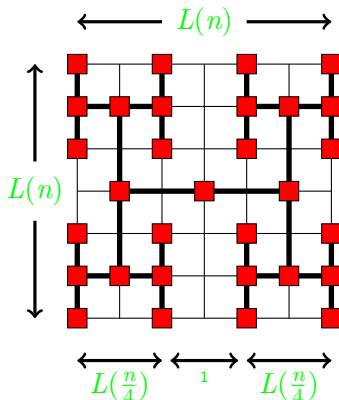
- Let's divide into 2 sub-trees, each with a size of $\frac{n}{2}$.



- We have:
$$H(n) = H\left(\frac{n}{2}\right) + 1 = \Theta(\log n)$$
$$W(n) = 2W\left(\frac{n}{2}\right) + 1 = \Theta(n)$$
- The area is $\Theta(n \log n)$.

Trial 2: divide into 4 sub-trees

- Let's divide into 4 sub-trees, each with a size of $\frac{n}{4}$.



- We have:
$$L(n) = 2L(\frac{n}{4}) + 1 = \Theta(\sqrt{n})$$
- Thus the area is $\Theta(n)$.