Divide and Conquer Algorithms

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- Solve each part recursively.
- Solve base cases by brute force.
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- Efficiently combine solutions for sub-problems into final solution.
- Common use:
  - Partition problem into two equal sub-problems of size $n/2$.
  - Solve each part recursively.
  - Combine the two solutions in $O(n)$ time.
  - Resulting running time is $O(n \log n)$.

Mergesort

SORT

INSTANCE: Nonempty list $L = x_1, x_2, \ldots, x_n$ of integers.

SOLUTION: A permutation $y_1, y_2, \ldots, y_n$ of $x_1, x_2, \ldots, x_n$ such that $y_i \leq y_{i+1}$, for all $1 \leq i < n$.

- Mergesort is a divide-and-conquer algorithm for sorting.
  1. Partition L into two lists A and B of size $\lfloor n/2 \rfloor$ and $\lceil n/2 \rceil$ respectively.
  2. Recursively sort A.
  3. Recursively sort B.
  4. Merge the sorted lists A and B into a single sorted list.

Merging Two Sorted Lists

- Merge two sorted lists $A = a_1, a_2, \ldots, a_k$ and $B = b_1, b_2, \ldots b_l$.

  Maintain a *current* pointer for each list.
  Initialise each pointer to the front of the list.
  While both lists are nonempty:

    Let a_i and b_j be the elements pointed to by the *current* pointers.
    Append the smaller of the two to the output list.
    Advance the current pointer in the list that the smaller element belonged to.

  EndWhile
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- Running time of this algorithm is $O(k + l)$.

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- Three basic ways of solving this recurrence relation:
  1. "Unroll" the recurrence (somewhat informal method).
  2. Guess a solution and substitute into recurrence to check.
  3. Guess solution in $O()$ form and substitute into recurrence to determine the constants.

Unrolling the recurrence

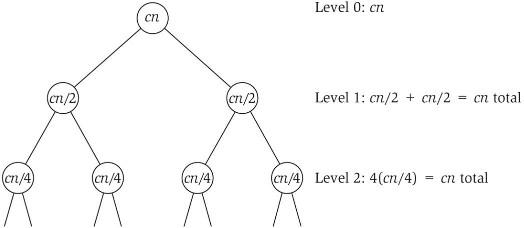

Figure 5.1 Unrolling the recurrence $T(n) \leq 2T(n/2) + O(n)$.

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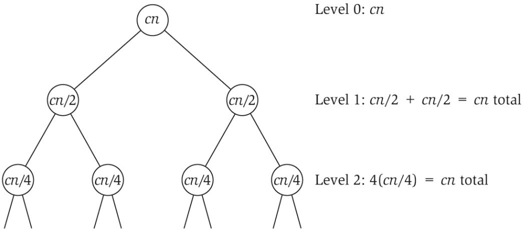

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- Recursion tree has $\log n$ levels.
- Total work done at each level is cn.
- Running time of the algorithm is $cn \log n$.
- Use this method only to get an idea of the solution.

Substituting a Solution into the Recurrence

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- Why is $T(n) \leq kn^2$ a "loose" bound?
- Why doesn't an attempt to prove $T(n) \leq kn$, for some $k > 0$ work?

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- Guess that the solution is $kn \log n$ (logarithm to the base 2).
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- $k \geq c$ will work.

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- Basic axiom: $T(n) \leq T(n+1)$, for all n: worst case running time increases as input size increases.
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- Let m be the smallest power of 2 larger than n.
- $T(n) \leq T(m) = O(m \log m) = O(n \log n)$, because $m \leq 2n$.

Other Recurrence Relations

- Divide into q sub-problems of size $n/2$ and merge in $O(n)$ time. Two distinct cases: $q = 1$ and $q > 2$.
- Divide into two sub-problems of size $n/2$ and merge in $O(n^2)$ time.

$T(n) = qT(n/2) + cn, q = 1$

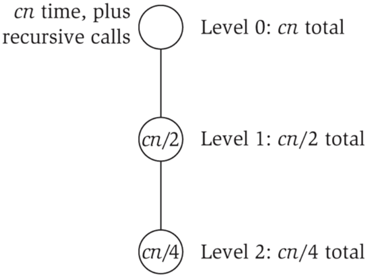

Figure 5.3 Unrolling the recurrence $T(n) \le T(n/2) + O(n)$.

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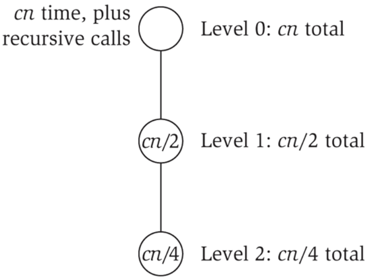

Figure 5.3 Unrolling the recurrence $T(n) \leq T(n/2) + O(n)$.

- Each invocation reduces the problem size by a factor of 2 $\Rightarrow$ there are $\log n$ levels in the recursion tree.
- At level i of the tree, the problem size is $n/2^i$ and the work done is $cn/2^i$.
- Therefore, the total work done is

$$\sum_{i=0}^{i=\log n} \frac{cn}{2^i} = O(n).$$

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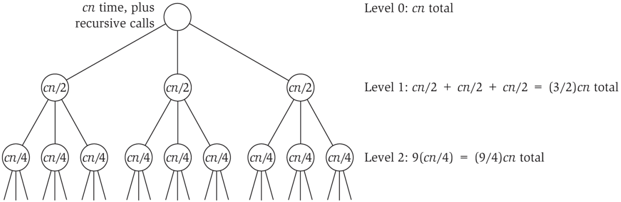

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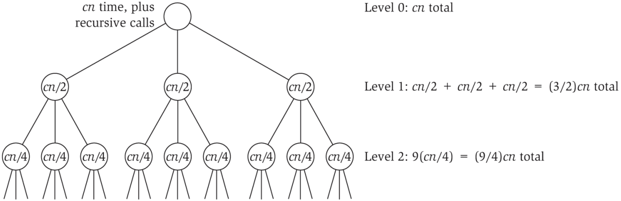

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- There are $\log n$ levels in the recursion tree.
- At level i of the tree, there are q^i sub-problems, each of size $n/2^i$.
- The total work done at level i is $q^i cn/2^i$. Therefore, the total work done is

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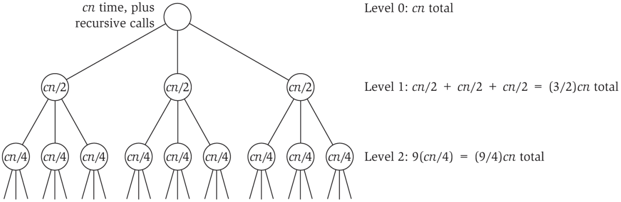

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$$\begin{aligned} T(n) &\leq \sum_{i=0}^{i=\log_2 n} q^i \frac{cn}{2^i} \leq cn \sum_{i=0}^{i=\log_2 n} \left(\frac{q}{2}\right)^i \\ &= O\left(cn \left(\frac{q}{2}\right)^{\log_2 n}\right) = O\left(cn \left(\frac{q}{2}\right)^{(\log_{q/2} n)(\log_2 q/2)}\right) \\ &= O(cn\, n^{\log_2 q/2}) = O(n^{\log_2 q}). \end{aligned}$$

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Computational Geometry

- Algorithms for geometric objects: points, lines, segments, triangles, spheres, polyhedra, ldots.
- Started in 1975 by Shamos and Hoey.
- Problems studied have applications in a vast number of fields: ecology, molecular biology, statistics, computational finance, computer graphics, computer vision, . . .

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INSTANCE: A set P of n points in the plane

SOLUTION: The pair of points in P that are the closest to each other.

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- At first glance, it seems any algorithm must take $\Omega(n^2)$ time.
- Shamos and Hoey figured out an ingenious $O(n \log n)$ divide and conquer algorithm.

Closest Pair: Set-up

- Let $P = \{p_1, p_2, \ldots, p_n\}$ with $p_i = (x_i, y_i)$.
- Use $d(p_i, p_j)$ to denote the Euclidean distance between p_i and p_j. For a specific pair of points, can compute $d(p_i, p_j)$ in $O(1)$ time.
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  - Divide and conquer after sorting: closest pair must be closest of
    1. closest pair in left half: distance δ_l.
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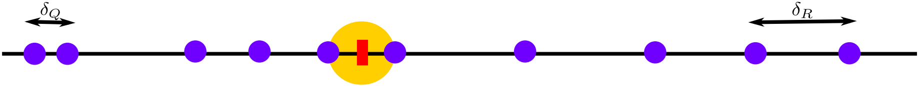

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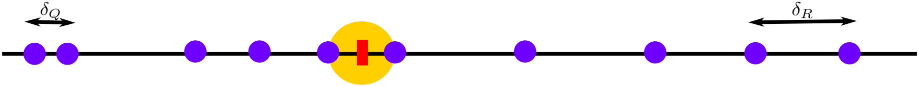
- Generalize the second idea to 2D.



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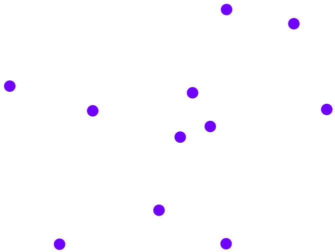

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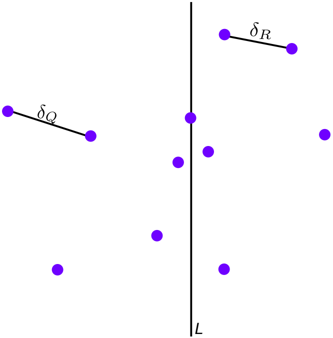

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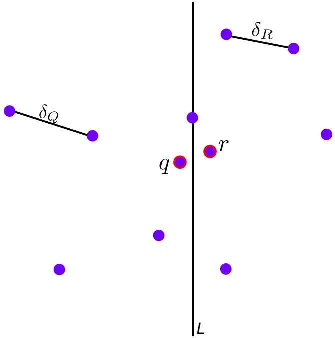
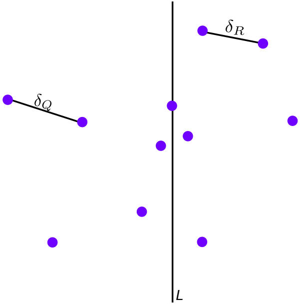
3. Let δ_Q be the distance computed for Q, δ_R be the distance computed for R, and $\delta = \min(\delta_Q, \delta_R)$.
4. Compute pair (q, r) of points such that $q \in Q$, $r \in R$, $d(q, r) < \delta$ and $d(q, r)$ is the smallest possible.



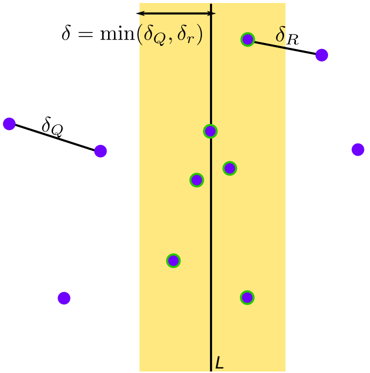
Closest Pair: Proof Sketch

- Prove by induction: Let (s, t) be the closest pair.
  1. both are in Q: computed correctly by recursive call.
  2. both are in R: computed correctly by recursive call.
  3. one is in Q and the other is in R: computed correctly in $O(n)$ time by the procedure we will discuss.
- Strategy: Pairs of points for which we do not compute the distance between cannot be the closest pair.
- Overall running time is $O(n \log n)$.



Closest Pair: Conquer Step

- Line L passes through right-most point in Q.
- Let S be the set of points within distance δ of L. (In image, $\delta = \delta_R$.)



Closest Pair: Conquer Step

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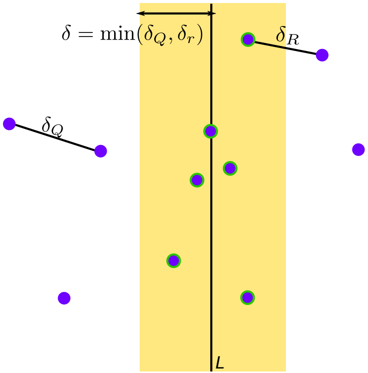

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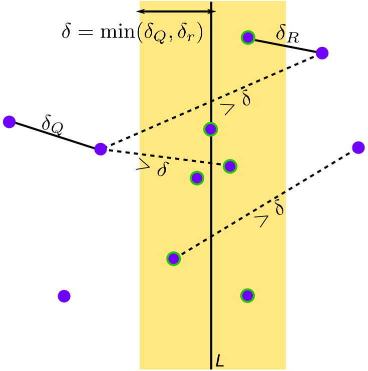
- Corollary: If $t \in Q - S$ or $u \in R - S$, then (t, u) cannot be the closest pair.



Closest Pair: Packing Argument

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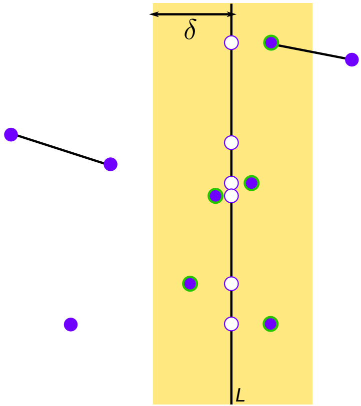

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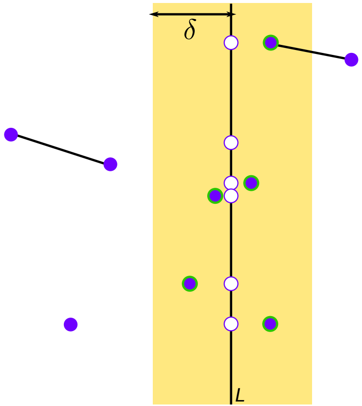
- Claim: If there exist $s, s' \in S$ such that $d(s, s') < \delta$ then s and s' are at most 15 indices apart in S_y.



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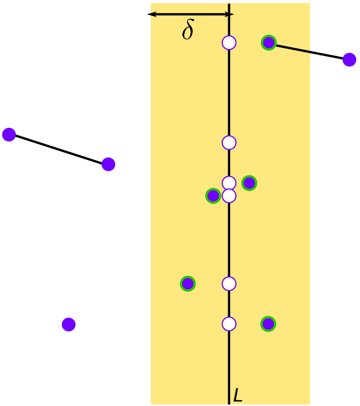

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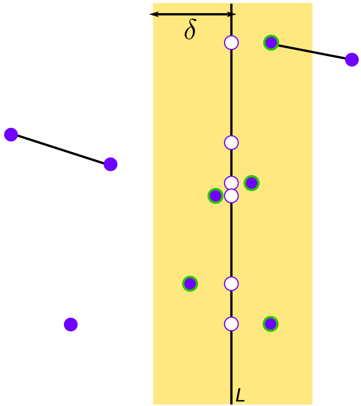
- Use the claim in the algorithm: For every point $s \in S_y$, compute distances only to the next 15 points in S_y.
- Other pairs of points cannot be candidates for the closest pair.



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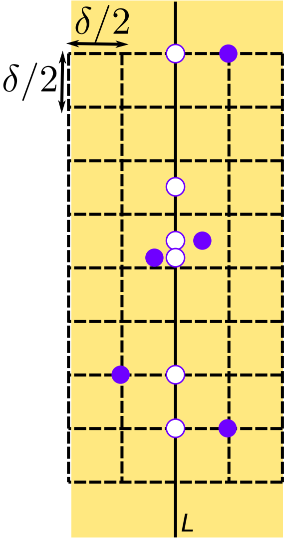

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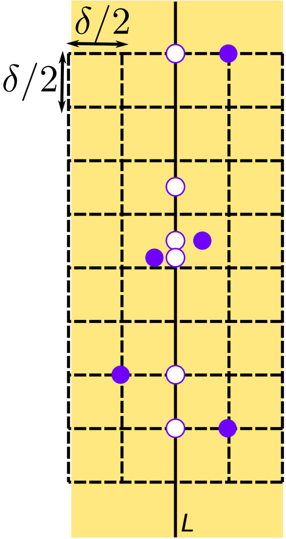
- Claim: If there exist $s, s' \in S$ such that s' appears 16 or more indices after s in S_y, then $s'_y - s_y \geq \delta$.
- Pack the plane with squares of side $\delta/2$.



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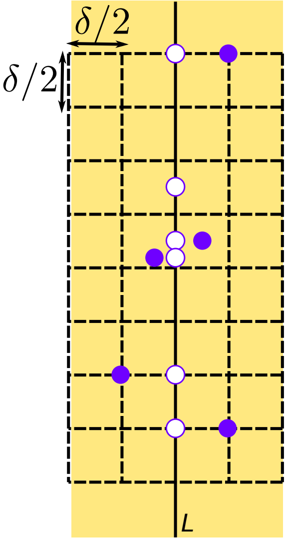

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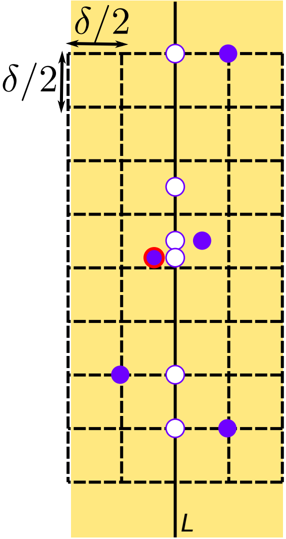

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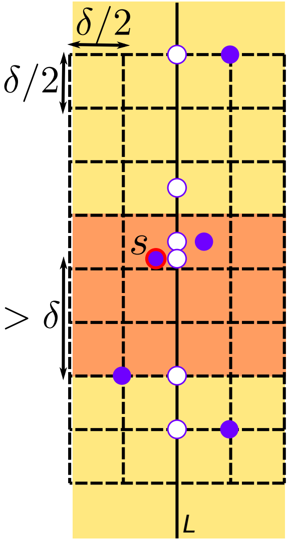
- Any point in the third row of the packing below s has a y-coordinate at least δ more than s_y.



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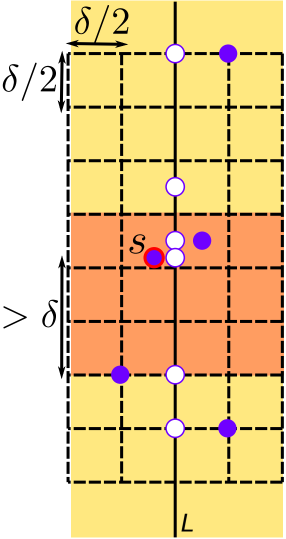
- Let s lie in one of the squares.
- Any point in the third row of the packing below s has a y-coordinate at least δ more than s_y.
- We get a count of 12 or more indices (textbook says 16).



Closest Pair: Final Algorithm

```
Closest-Pair(P)
  Construct P_x and P_y  (O(n log n) time)
  (p*_0, p*_1) = Closest-Pair-Rec(P_x,P_y)

Closest-Pair-Rec(P_x, P_y)
  If |P| ≤ 3 then
    find closest pair by measuring all pairwise distances
  Endif

  Construct Q_x, Q_y, R_x, R_y (O(n) time)
  (q*_0,q*_1) = Closest-Pair-Rec(Q_x, Q_y)
  (r*_0,r*_1) = Closest-Pair-Rec(R_x, R_y)

  δ = min(d(q*_0,q*_1), d(r*_0,r*_1))
  x* = maximum x-coordinate of a point in set Q
  L = {(x,y) : x = x*}
  S = points in P within distance δ of L.

  Construct S_y (O(n) time)
  For each point s ∈ S_y, compute distance from s
     to each of next 15 points in S_y
     Let s, s' be pair achieving minimum of these distances
     (O(n) time)

  If d(s,s') < δ then
     Return (s,s')
  Else if d(q*_0,q*_1) < d(r*_0,r*_1) then
     Return (q*_0,q*_1)
  Else
     Return (r*_0,r*_1)
  Endif
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($O(n)$ time)

If $d(s,s') < \delta$ then

Return (s,s')

Else if $d(q_0^*,q_1^*) < d(r_0^*,r_1^*)$ then

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Else

Return (r_0^*,r_1^*)