Theory of Parallel and Distributed Systems (WS2016/17)
Chapter 3
More Algorithms

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Colorings

Coloring Problem

- Given undirected graph $G = (V, E)$ and $k \in \mathbb{N}$.
- Compute [exists?] Function $c : V \mapsto \{1, \cdots, k\}$ with:
  - $\forall \{a, b\} \in E : c(a) \neq c(b)$.
- Coloring number (chromatic index) of $G$:
  $$\chi(G) := \min\{\{k \mid \exists c : V \mapsto \{1, \cdots, k\} \mid \forall \{a, b\} \in E : c(a) \neq c(b)\}\}.$$

- Coloring problem is NP-complete.
- Let $G = C_n$, i.e. $G = (\{v_0, \cdots, v_{n-1}\}, \{\{v_i, v_{(i+1) \mod n}\} \mid 0 \leq i < n\})$.
- Then we have $\chi(C_n) \leq 3$ and $\chi(C_{2 \cdot n}) \leq 2$ ($\chi(C_{2 \cdot n+1}) = 3$).
- We do not have a nice order on the nodes:
- Let $\pi(i)$ be a permutation
- Let $G = C_n$, i.e.
  $$G = (\{v_0, \cdots, v_{n-1}\}, \{\{v_{\pi(i)}, v_{\pi((i+1) \mod n)}\} \mid 0 \leq i < n\}).$$
Parallel Coloring Algorithm of (on) a cycle (Idea)

- A processor $P_i$ works on $v_{\pi(i-1)}$ for some permutation $\pi$.
- Register $R_i$ holds $\pi(i - 1)$.
- Register $N_i$ holds $\pi(i)$.
- In register $C_i$ will be the color of $v_{R_i}$.
- Initialize $C_i$ with $i$.
- Reduce step by step the number of colors.
- We will use the colors $\{0, 1, \cdots, n\}$. 
Parallel Coloring Algorithm of (on) a cycle (Idea)

Programm: color-cycle

for all $P_{i+1}$ where $0 \leq i < n$ do in parallel

$\pi(i - 1) \rightarrow R_i$
$\pi(i) \rightarrow N_i$
$c = i$
$c \rightarrow C_i$

repeat $\lceil \log^*(n) \rceil + 2$ times

$C_{N_i} \rightarrow c'$
minimal $k$ with: $((c \gg k) \% 2) \neq ((c' \gg k) \% 2)$.
$c = 2 \cdot k + ((c \gg k) \% 2)$.
$c \rightarrow C_i$
Parallel Coloring Algorithm of (on) a cycle (Idea)

- At the start we are using $n$ colors.
- Within each color-reduction will the coloring stay correct.
- Within each color reduction will the coloring number be reduced from $x$ to $\log(x) + O(1)$.
- After $\lceil \log^*(n) \rceil$ reductions steps will be the coloring numbers $\leq 5$.
- A second reduction of colors will follow now:
The rows hold $c$ and the columns hold $c'$.  

The entries in the table hold the new $c$.

\[
\begin{array}{cccccccc}
& 000 & 001 & 010 & 011 & 100 & 101 & 110 & 111 \\
0 & 000 & 0 & 2 & 0 & 4 & 0 & 2 & 0 \\
1 & 001 & 1 & 1 & 2 & 1 & 4 & 1 & 2 \\
2 & 010 & 3 & 0 & 0 & 3 & 0 & 4 & 0 \\
3 & 011 & 1 & 3 & 1 & 1 & 3 & 1 & 4 \\
4 & 100 & 5 & 0 & 2 & 0 & 0 & 2 & 0 \\
5 & 101 & 1 & 5 & 1 & 2 & 1 & 1 & 2 \\
6 & 110 & 3 & 0 & 5 & 0 & 3 & 0 & 0 \\
7 & 111 & 1 & 3 & 1 & 5 & 1 & 3 & 1 \\
\end{array}
\]

We only have the colors 000,001,010,011,100,101 ($\leq 5$).
Parallel Coloring Algorithm of (on) a cycle (Idea)

Programm: color-cycle

for all $P_{i+1}$ where $0 \leq i < n$ do in parallel

$\pi(i-1) \to R_i$
$\pi(i) \to N_i$
$c = i$
$c \to C_i$

repeat $\lceil \log^* (n) \rceil + 2$ times

$C_{N_i} \to c'$

minimal $k$ with: $((c \gg k) \% 2) \neq ((c' \gg k) \% 2)$.
$c = 2 \cdot k + ((c \gg k) \% 2)$.
$c \to C_i$

for $r := 5$ downto $3$ do:

if $c = r$ then

$C_{N_i} \to c'$
$c' \to C_i$
$C_{N_i} \to c''$
$c := \min(\{0, 1, 2\} \setminus \{c', c''\})$
$c \to C_i$
Coloring a Cycle

**Theorem:**
A cycle with \( n \) nodes could be colored with \( n \) processors in time \( O(\log^* n) \) with at most 3 colors.

Proof: see above.

**Theorem:**
A cycle of \( n \) processors may color itself in time \( O(\log^* n) \) with at most 3 colors.

Proof: see above.

**Theorem:**
A cycle of \( n \) processors needs at least \( \Omega(\log^* n) \) time to color itself with at most 3 colors.

Proof: see V4.
Coloring a Tree

- A processor $P_i$ works on $v_{\pi(i-1)}$ for some permutation $\pi$.
- Register $R_i$ holds $\pi(i - 1)$.
- Register $N_i$ holds $\pi(j - 1)$ where $j$ is the father of $i$.
- The father of the root $r$ is $r$.
- In register $C_i$ will be the color of $v_{R_i}$.
- Initialize $C_i$ with $i$.
- Reduce step by step the number of colors.
- We will use the colors $\{0, 1, \cdots, n\}$. 
Parallel Coloring Algorithm of (on) a tree (Idea)

Programm: color-cycle

for all $P_{i+1}$ where $0 \leq i < n$ do in parallel

\[\pi(i - 1) \rightarrow R_i\]
\[\pi(i) \rightarrow N_i\]
\[c = i\]
\[c \rightarrow C_i\]

repeat $\lceil \log^*(n) \rceil + 2$ times

\[C_{N_i} \rightarrow c'\]
minimal $k$ with: $((c \gg k) \% 2) \neq ((c' \gg k) \% 2)$.
\[c = 2 \cdot k + ((c \gg k) \% 2)\]
\[c \rightarrow C_i\]
Parallel Coloring Algorithm of (on) a tree (Idea)

- At the start we are using $n$ colors.
- Within each color-reduction will the coloring stay correct.
- Within each color reduction will the coloring number be reduced from $x$ to $\log(x) + O(1)$.
- After $\lceil \log^*(n) \rceil$ reductions steps will be the coloring numbers $\leq 5$.
- A second reduction of colors will follow now:
Parallel Coloring Algorithm of (on) a tree (Idea)

Programm: color-tree

for all $P_{i+1}$ where $0 \leq i < n$ do in parallel

$\pi(i-1) \rightarrow R_i$
$\pi(i) \rightarrow N_i$
$c = i$ and $c \rightarrow C_i$

repeat $\lceil \log^*(n) \rceil + 2$ times, if $R_i \neq N_i$

$C_{N_i} \rightarrow c'$
minimal $k$ with: $((c \gg k) \% 2) \neq ((c' \gg k) \% 2)$.
$c = 2 \cdot k + ((c \gg k) \% 2)$.
$c \rightarrow C_i$

for $r := 5$ downto 3 do:

if $c = r$ then

$C_{N_i} \rightarrow c'$
$c' \rightarrow C_i$
$C_{N_i} \rightarrow c''$
$c := \min(\{0, 1, 2\} \setminus \{c', c''\})$
$c \rightarrow C_i$
Parallel Coloring Algorithm of (on) a tree (Idea)

Programm: color-cycle

for all $P_{i+1}$ where $0 \leq i < n$ do in parallel

\begin{align*}
\pi(i - 1) &\rightarrow R_i \\
\pi(j - 1) &\rightarrow N_i \text{ with } j \text{ is father of } i \\
c &\equiv i \text{ and } c \rightarrow C_i
\end{align*}

repeat $\lceil \log^*(n) \rceil + 2$ times

\begin{align*}
C_{N_i} &\rightarrow c' \\
\text{minimal } k \text{ with: } ((c \gg k) \bmod 2) \neq ((c' \gg k) \bmod 2). \\
c &\equiv 2 \cdot k + ((c \gg k) \bmod 2). \\
c &\rightarrow C_i \\
\text{if } R_i = N_i \text{ then } c = \min(\{0, 1\} \setminus R_i) \text{ else } c = C_{N_i} \\
c &\rightarrow C_i
\end{align*}

for $r := 5$ downto 3 do:

if $c = r$ then

\begin{align*}
C_{N_i} &\rightarrow c' \\
c' &\rightarrow C_i \\
C_{N_i} &\rightarrow c'' \\
c &\equiv \min(\{0, 1, 2\} \setminus \{c', c''\}) \\
c &\rightarrow C_i
\end{align*}
Coloring a Tree

**Theorem:**
A tree with $n$ nodes could be colored with $n$ processors in time $O(\log^* n)$ with at most 3 colors.

Proof: see above.

**Theorem:**
A tree of $n$ processors may color itself in time $O(\log^* n)$ with at most 3 colors.

Proof: see above.
Eulerian cycle

Definition:
A graph \( G = (V, E) \) is called Eulerian, iff there exists a cycle which visits each edge precisely once.

Theorem
A non-directed graph \( G = (V, E) \) is Eulerian
- \( G \) is connected and
- each node of \( G \) has even degree.

Theorem
A directed graph \( G = (V, E) \) is Eulerian
- \( G \) is strongly connected and
- each node as as many incoming edges as outgoing ones.

Problem: Compute Eulerian cycle on Eulerian graphs.
Idea

- Non Parallel:
  - Start with a free edge and follow free/unused edges till a cycle is closed.
  - Repeat till all edges are in some cycle.
  - Join pairs of cycles into a single one.
  - Repeat till just one cycle remains.

- If $G$ is non-directed, then make a directed version of $G$.
- Compute a cover of cycles.
- Compute an additional cycle which meets each cycle precisely once.
- Uses these to compute a cycle for $G$
- Delete some edges to get a Eulerian cycle for $G$.
Change a non-directed Graph into a directed one

- $G$ contain $m$ non-directed edges.
- Substitute each non-directed edge with two directed ones: $\{i, j\}$ becomes $(i, j)$ and $(j, i)$.
- Define a successor for each edge:
  - The neighbors of $v$ are: $v_0, v_1, \cdots, v_{d-1}$.
  - Then define for all $i$:
    
    $Succ((v_i, v)) := (v, v_{(i+1) \mod d})$ und
    
    $Succ((v_{(i+1) \mod d}, v)) := (v, v_i)$.
- Each directed edge is in precisely one cycle (defined by $Succ$).
- For each cycle $C$ exists one cycle $C'$, which consists the reverse edges.
- We will now delete one of the two cycles $C$ or $C'$.
Identify the generated cycles:

- Let \( \text{min}(((i, j), (k, l))) := \begin{cases} (i, j) & \text{if } i \leq k \lor i = k \land j < l \\ (k, l) & \text{otherwise} \end{cases} \).
- For each edge \( e \) define \( \text{Edge}'(e) = e \);
- For all edges \( e \) repeat \( \log m \) times:
  - \( \text{Edge}'(e) = \min(\text{Edge}'(e), \text{Edge}'(\text{Succ}(e))) \)
  - \( \text{Succ}(e) = \text{Succ}(\text{Succ}(e)) \).
- For each edge \( (i, j) \): if \( \text{min}(((i, j), (j, i))) \neq (i, j) \) then let \( \text{Edge}'(e) = 0 \).
- Thus we have selected for each non-directed edge a directed one (resp. a direction).
- Possible with \( m \) in time \( O(\log m) \).
- We consider in the following on directed graphs.
Step 1

- Let $G = (V, E)$ be a directed graph.
- Sort the edges into an array $Edge$.
  - using the order: $(i, j) < (k, l) \iff j < l \lor (j = l \land i < k)$.
- Sort the edges into an array $Succ$.
  - using the order: $(i, j) < (k, l) \iff i < k \lor (i = k \land j < l)$.
- We have already defined the cycles:
  - Successor of edge $e = Edge(i)$ is the edge $Succ(i)$.
- We also store in $P(i)$ the position of $Succ(i)$ in $Edge$.
- I.e. $Edge(P(i)) = Succ(i)$.
- This information could be updated during the sorting of $Succ$.
- This could be done in time $O(\log m)$ using $O(m)$ processors.
Step 2

- Situation: We have a directed graph covered by cycles.
- Problem: Compute for each edge \( e \) the cycles where \( e \) belongs to.
- Solution: compute for each cycle the minimal edge 

\[
((i, j) < (k, l) \iff i < k \lor (i = k \land j < l))
\]

- Algorithm:

Programm:

\[
\text{for all } P_i \text{ where } 1 \leq i \leq m \text{ do in parallel}
\]

\[
\text{CycleRep}(i) := \text{Succ}(i)
\]

\[
\text{for } i := 1 \text{ to } \lceil \log m \rceil \text{ do:}
\]

\[
\text{CycleRep}(i) := \min(\text{CycleRep}(i), \text{CycleRep}(P(i)))
\]

\[
P(i) := P(P(i))
\]

- We use again the doubling technique.
- Possible in time \( O(\log m) \) using \( O(m) \) Processors.
Step 2 (Continued)

- Situation: the cycles of the coverage are identified by \( \text{CycleRep} \).
- Problem: join the cycle into a single one.
- Solution: Identify the nodes of the cycle.
  
  \[ C = \{ \text{CycleRep}(i) | 1 \leq i \leq m \} \text{. (Note } C \text{ is a edge set)} \]

- \( G' = V \cup C \)

- \( E' = \{(u, v) | u \in V, v \in C : v \text{ is identified in the cycle by } u\} \)

- Computing of \( E' \):

  **Programm:**

  ```
  for all \( P_i \) where \( 1 \leq i \leq m \) do in parallel
  (u, v) = \text{Edge}(i)
  \text{Edge}'(2 \cdot i) = (u, \text{CycleRep}(i))
  \text{Edge}'(2 \cdot i + 1) = (v, \text{CycleRep}(i))
  ```
Step 2 (Continued)

- **Situation:** Cover of cycles and graph $G'$ defined.
- **Problem:** there are multiple edges.
- **Solution:** sort them out.
- **Sort** $Edge'$.
- **Programm:**
  
  ```
  for all $P_i$ where $1 \leq i \leq m$ do in parallel 
  if $Edge'(i) = Edge'(i + 1)$ then $Edge(i) = \infty$
  ```
  
- **Sort** $Edge'$.
- **Consider only the first** $|E'|$ elements of $Edge'$.
- **Problem:** node $u$ could appear several times in a cycle $v$.
- **As before we may compute a single representative.**
- **Let these edge be** $(i, u) = Cert(u, v)$.
- **May be done in time** $O(\log m)$ using $O(m)$ processors.
Step 3

- Situation: Covering of the cycles and graph $G'$ computed.
- Problem: Compute cycle in $G'$.
- Solution: compute spanning tree $T$ for the bipartite Graph $G'$.
  To compute spanning tree we need $O(\log^2 m)$ time with $O(m/\log^2 m)$ Processors.
- Then we substitute each edge in $T$ with two directed edges.
- The new graph $T'$ is Eulerian.
- The Eulerian cycle is easy to find:
  To do so, compute for each node of the tree the order of edges.
  Could be done in time $O(\log m)$ using $O(m)$ processors.
Step 4

- Situation: We have a cover of cycles for $G$ and $T'$. 
- Problem: Find cycle $L$ in $G'$. 
- Solution: Combine the cycles using $\text{Cert}(u, v)$. 
- $L$ will also contain the Eulerian cycle in $G$. 
- For each cycle $v$ in $G$ $\text{Cert}(u, v)$ gives us an edge, at which we may exchange between $v$ and the cycle in $T'$. 
- These points of change will be used to construct a single cycle $L$. 
- Time $O(1)$ using $O(m)$ Processors.
Step 5

- Situation: we have a cycle for $G$ and $T'$.
- Problem: find cycle in $G$.
- Solution: delete edges from $T'$.
- Programm:
  
  ```
  for all $P_i$ where $1 \leq i \leq m$ do in parallel
  if $Succ(i) \in T'$ then $Succ(i) := Succ(Succ(i))$
  if $Succ(i) \in T'$ then $Succ(i) := Succ(Succ(i))$
  ```

- Uses time $O(1)$ with $O(m)$ processors.
- Total time is: $O(\log^2 m)$ using $O(m)$ processors.
- Also possible: $O(\log^2 m)$ time using $O(m/\log^2 m)$ processors.
Definition

Let $G = (V, E)$ be a non-directed graph.

- $M \subseteq E$ is called a matching, iff $\forall e, e' \in M : e \cap e' = \emptyset$.
- $M$ is called maximal matching, iff $\not\exists e \in E : M \cup \{e\}$ is a matching.
- $M$ is called maximum matching, iff for all matchings $M'$ we have $|M'| \leq |M|$.

Sequential: $O(m \log m)$ for maximal matching.

Idea: Choose any free edge and delete all incident edges.

Sequential: $O(m^3)$ for maximum matching.

Idea: enlarging alternating pathes.
Let $\Delta(G)$ be the maximal degree of $G$.

- Enlarge the matching step by step by several edges.
- There will be $O(\log_{3/2} n)$ phases.
- $i$-te phase $F_i$ has $G_i$ as input and will output $M_i$.
- $G_1 = G$ and final result: $\bigcup M_i$.
- Within each phase $F_i$ we will call the procedure $DegreeSplit$ $(1 + \log(\Delta(G)))$-times.
- Within each step within a phase we will half the node degree.
- We denote with $G(i,j)$ the graph considered in the $j$-th Step of the $i$-th phase.
- We will describe the procedure $DegreeSplit$.
- Let $k$ be the smallest number with $2^k \leq \Delta(G) \leq 2^{k+1}$.
- We will call all nodes $v$ with $\delta(v) \geq 2^k$ active.
Step 1

- Compute all active nodes of $G(i,j)$
  - Determine the degree in time $O(\log \Delta(G(i,j)))$ with $O(m)$ processors.
  - Determine the maximum degree in time $O(\log n)$ with $O(n)$ processors.
  - Then the active nodes are known in time $O(1)$ using $O(n)$ processors.

- Total running time: $O(\log n)$ using $O(m)$ processors.
Step 2

- Compute the graph $G^*(i, j)$ as follows:
  - Compute all nodes that are incident to active nodes.
  - Determine the new node degree.
  - If there are nodes with odd degree connect them to a new node $v$.

- Total running time: $O(\log n)$ using $O(m)$ processors.
- $G^*(i, j)$ might not be connected.
- Each component of $G^*(i, j)$ contains an Eulerian cycle.
- Note that each node $v$ has even degree.
Step 3

- Compute an Eulerian cycle on each component of $G^*(i,j)$.
- This needs time $O(\log^2 n)$ with $O(m + n)$ processors.
- Note that the additional $n$ processors result from the additional edges.
- Label the edges from the Eulerian cycle alternating with 0 and 1.
- For the component with the additional node $v$ start with $v$ using label 0.
- For all other components start at an arbitrary node with label 1.
- Running time: $O(\log n)$ with $O(m + n)$ processors.
- Use Parallel Prefix to compute the labels.
Step 4

- Delete all edges with label 0.
- If the remaining graph $G^{**}(i,j)$ is not a matching then $G(i,j + 1) = G^{**}(i,j) \setminus \{v\}$.
- If the remaining graph $G^{**}(i,j)$ is a matching then $M_i = E(G^{**}(i,j))$.
- Running time: $O(1)$ with $O(m + n)$ processors.
- Running time of the procedure DegreeSplit: $O(\log^2 n)$ with $O(m + n)$ processors.
- It remains to show: After at most $1 + \log(\Delta(G(i,j)))$ steps DegreeSplit computes a matching.
- It remains to show: After at most $O(\log_{3/2} n)$ phases the matching is optimal.
Inner loop

**Lemma:**
Let $G$ be the input of $DegreeSplit$, then $DegreeSplit$ will compute a matching after $1 + \log(\Delta(G))$ iterations.

**Proof:**
- Let $k$ be the smallest number with $2^k \leq \Delta(G) \leq 2^{k+1} + 1$.
- Let $G_1$ be the result of an iteration.
- Let $v$ be active in $G$. It holds:
  - $2^k \leq \delta_G(v)$.
  - $\lfloor \delta_G(v)/2 \rfloor \leq \delta_{G_1}(v) \leq \lceil \delta_G(v)/2 \rceil + 1$.
  - $2^{k-1} \leq \delta_{G_1}(v) \leq 2^k + 1$.
- Then $v$ stays active in $G_1$.
- Hence the degree is halved in every step.
- There exists a $k' \leq k$ such that $G_{k'}$ has a degree of 3.
- After two more iterations the degree is at most one.
- So a matching is found.
Lemma:
A logarithmic number of phases is enough to compute a maximum matching.

Proof:
- Let $A_i$ be the nodes that are active in phase $F_i$.
- Then $A_i$ is a vertex cover of $G_i$.
- ($C \subseteq V$ is a vertex cover if $\forall e \in E : e \cap C \neq \emptyset$)
- We show the following;
  - Half of the nodes in a vertex cover $A_i$ can be made incident to edges from $M_i$.
  - This means it holds: $|A_i/G_{i+1}| \leq |A_i|/2$.
    with $A_i/G_{i+1}$ the nodes of $A_i$ in $G_{i+1}$
  - There are vertex covers $C_i$: $|C_{i+1}| \leq 2 \cdot |C_i|/3$. 

Outer loop
Outer loop (Proof)

- Let $G_k = (V, E_k)$ be the graph in the third to last loop of $DegreeSplit$.
- W.l.o.g. $G_k$ is connected with degree $\leq 3$.
- $DegreeSplit$ can w.l.o.g. remove the smallest set of edges.
- Hence it holds $|M_i| \geq |E_k|/4$. 
Outer loop (Proof)

- If \(|E_k| \geq |A_i|\) then \(M_i\) contains at least \(|A_i|/4\) edges.
  - Both end points of an edge from \(A_i\) belong to \(A_i\) and
  - at least half of them are incident to \(M_i\).

- If \(|E_k| < |A_i|\) then \(G_k\) is a tree.
  - We remove edges from \(G_k\) that have a leaf as one of its end points.
  - Furthermore the incident edges are removed.
Outer loop (Proof)

- Because $\Delta(G_k) \leq 3$ at most 2 trees $T_1$ and $T_2$ remain (with $n_1 + n_2$ nodes).
- Then $((n_1 - 1) + (n_2 - 1))/4$ edges are added to $M_i$.
- Then $M_i$ contains $|A_i|/2$ nodes.
- Then it holds: $|A_i/G_i+1| \leq |A_i|/2$. 
We show using induction that $G_i$ contains a vertex cover $C_i$ with $|C_i| \leq (2/3)^{i-1}|V|$.

We will show that $|C_{i+1}| \leq 2|C_i|/3$.

**Basis**: $i = 1$: Choose $C_1 = V$.

**Case 1**: $|A_i| \leq 4|C_i|/3$.
- In phase $i$ half of the nodes are removed from $A_i$.
- $A_i/G_{i+1}$ is a vertex cover from $G_{i+1}$.
- $|A_i|/2 \leq (4|C_i|/3)/2 = 2|C_i|/3$.

**Case 2**: $|A_i| > 4|C_i|/3$.
- Half of the nodes from $A_i$ are removed.
- These have end points in $M_i$.
- $C_i$ is a vertex cover of $G_i$.
- Then every edge has at least one end point in $C_i$.
- At least $1/4$ of the edges in $A_i$ are contained in $C_i$.
- $C_i/G_{i+1}$ is a vertex cover of $G_{i+1}$.
- $|C_i/G_{i+1}| \leq |C_i| - |A_i|/4 \leq |C_i| - (4|C_i|/3)/4 = 2|C_i|/3$. 

$|M_i| \geq |E_k|/4$ 
$|A_i/G_{i+1}| \leq |A_i|/2$
Summary

Theorem:

A maximal vertex cover can be computed in time $O(\log^4 n)$ using $O(n + m)$ processors.

Proof:

- Outer loop: $O(\log n)$
- Inner loop: $O(\log n)$
- Running time of $\text{DegreeSplit}$: $O(\log^2 n)$.

\[
|M_i| \geq |E_k|/4 \\
A_i / G_{i+1} \leq |A_i|/2
\]
Bipartite graphs

A graph \( G = (A, B, E) \) with \( E \subset A \times B \) is called bipartite graph.

directed line graph

Let \( G = (V, E) \) be a directed graph, then \( G^2 = (E, F) \) with
\[
F = \{((a, b), (b, c)) \mid (a, b), (b, c) \in E\}
\]
is the line graph of \( G \).

undirected line graph

Let \( G = (V, E) \) be an undirected graph, then \( G^2 = (E, F) \) with
\[
F = \{\{\{a, b\}, \{b, c\}\} \mid \{a, b\}, \{b, c\} \in E\}
\]
is the line graph of \( G \).

Edge coloring

- Let \( G = (V, E) \) be an undirected graph and \( k \in \mathbb{N} \).
- Compute \( [\exists \text{ a } k\text{-coloring of } G^2] \).
Introduction

- Let $G = (V, E)$ be an undirected graph.
- It is NP-complete to find a $\Delta(G)$ edge coloring.
- There is always a $\Delta(G) + 1$ edge coloring.
- A bipartite graph $G$ is $\Delta(G)$ edge colorable.
- Or: A bipartite graph $G$ can be covered with $\Delta(G)$ matchings.
- Here: Parallel edge coloring of a bipartite graph.
- 1.Step: $\Delta(G) = 2^k$ for some $k \in \mathbb{N}$. 

3.39 Bipartite graphs
Method for $\Delta(G) = 2^k$

- Idea: Cover the edges of $G$ with cycles and paths.
- Color edges alternating with 0 and 1.
- This computes a partition of $G$ in $G_0$ and $G_1$ with $\Delta(G_0) = \Delta(G_1) = 2^{k-1}$.
- All steps can be done in time $O(\log n)$ with $O(m)$ processors.
- Continue recursively.
- Total running time: $O(\log^2 n)$ with $O(m)$ processors.
Example
Example
Example
Method for $\Delta(G) < 2^k$ (Idea)

- Color as many edges as possible in the sub graph $G'$ with $\Delta(G') = 2^{k'}$.
- Allow double coloring of edges, i.e. $(i, j)$ is colored $\alpha$ at $i$ and $\beta$ at $j$.
- Within each step it holds:
  - There are correctly colored edges and double colored edges.
  - The set of colors $S$ is chosen such that the number of double colored edges is as big as possible,
  - These edges become colored correctly.
  - This happens in the extended sub graph with $\Delta(G') = 2^{k'}$. 
Method for $\Delta(G) < 2^k$ (Idea)

- Let $k': 2^{k'} < \Delta(G) < 2^{k'+1}$, $C = \emptyset$ and $U = E$.
- Partition $F = \{0, 1, 2, \ldots, \Delta(G) - 1\}$ into four sets of almost the same size $S_1, S_2, S_3, S_4$.
- Repeat until all edges are colored correctly:
  - Choose double coloring of the edges from $U$.
  - Chose $i, j$ with: As many edges as possible from $U$ are colored with only $S_i \cup S_j$.
  - Let $U'$ be those edges.
  - It holds: $|U'| \geq |U|/6$ and $U' \leq 2^{k'}$.
  - Let $H$ be those edges that only use colors from $S_i \cup S_j$.
  - Let $G' = (V, H)$, extend $G'$ such that $\Delta(G') = 2^{k'}$.
  - Color $G'$ using the method from above.
  - Set $C = C \cup H$, these are the correctly colored edges.
- Total running time: $O(\log^3 n)$ with $O(m)$ processors.
Example (1. round)
Example (2. round)
Example (2. round)
Example (3. round)
Example (3. round)
Example (result)
Results

Theorem:
A bipartite graph $G$ with $\Delta(G) = 2^k$ can be edge colored with $\Delta(G)$ colors in time $O(\log^2 n)$ with $O(m)$ processors.

Proof: See above.

Theorem:
A bipartite graph $G$ can be edge colored with $\Delta(G)$ colors in time $O(\log^3 n)$ with $O(m)$ processors.

Proof: See above.
Results without proof

**Lemma**

Any graph $G = (V, E)$ with maximal degree $\Delta$ is $\Delta + 1$ colorable.

**Lemma**

Any graph $G = (V, E)$, which is not a clique nor a odd cycle is $\Delta$ colorable.

- Idea of distributed/parallel algorithm:
  - Reduce recursively the colors.
  - Double the size of correctly colored sub-graphs.
  - Or use the idea for trees to bounded degree graphs.
Recall and Idea 1

**Theorem:**
A tree with \( n \) nodes could be colored with \( n \) processors in time \( O(\log^* n) \) with at most 3 colors.

- Recall: choose minimal \( k \) with: \((c \gg k)\%2) \neq ((c' \gg k)\%2)\) and
- set \( c = 2 \cdot k + ((c \gg k)\%2)\).
- This did produce a 6-coloring on trees.
- On a bounded degree graph use this idea on a vector of length \( \Delta \).
Algorithm 1

choose minimal $k$ with: $((c \gg k)\%2) \neq ((c' \gg k)\%2)$ and set $c = 2 \cdot k + ((c \gg k)\%2)$

1. Let $v_1, v_2, \ldots, v_d$ the $d \leq \Delta$ neighbors of $v$
2. Let $c_1, c_2, \ldots, c_d$ the colors $v_i$ and $c$ the color of $v$.
3. For each $i$ ($1 \leq i \leq d$) do
   1. choose minimal $k_i$ with: $((c \gg k_i)\%2) \neq ((c_i \gg k_i)\%2)$ and
   2. set $b_i = 2 \cdot k_i + ((c \gg k_i)\%2)$.
4. Choose new color for $v$: $(b_1, b_2, \ldots, b_d)$.

- As before, the coloring stays valid.
- Like before, a $x$-bit coloring becomes a $\Delta(\log x + 1)$-bit coloring.
- Like before, we may reduce the colors to $\Delta + 1$ colors.
- For unbounded degree the running time becomes: $O(\log^* n + 2^\Delta)$. 
Theorems 1

Choose minimal $k$ with: $((c \gg k) \mod 2) \neq ((c' \gg k) \mod 2)$ and set $c = 2 \cdot k + ((c \gg k) \mod 2)$.

Theorem

A constant degree graph may be colored with $\Delta + 1$ colors in time $O(\log^* n)$ on a distributed system.

Theorem

A constant degree graph may be colored with $\Delta + 1$ colors in time $O(\log^* n)$ on a parallel system using $n$ processors.
Notations and Idea 2

choose minimal $k$ with: $((c \gg k)\%2) \neq ((c' \gg k)\%2)$ and set $c = 2 \cdot k + ((c \gg k)\%2)$

- $x$ will be a binary string with up to $k$ bits.
- Define $U_x = \{(a_1, a_2, ... a_{k-|x|}, x) \mid a_i \in \{0, 1\}\}$.
- The procedure RecurseColor will color $U_x$ with $\Delta + 1$ colors.
- Idea:
  - Having colored $U_x$ with $\Delta + 1$ colors,
  - Recolor $U_{1x}$ such that $U_{0x}$ and $U_{1x}$ are colored correctly.
  - This doubles the size of correctly colored sub-graphs.
Recursive Algorithm

RecurseColor(x) (initial with x = ε):

1. Let ID = (a1, a2, ..., ak) be a vector of bits, which identify the node/prozessor v.
2. Set l = |x|.
3. If l = k then set c(v) = 1 and return.
4. Set b = ak−l.
5. Set c(v) = RecurseColor(bx).
6. If b = 0 then return.
7. For round i from 1 to Δ + 1 do
   a. if c(v) = i then c(v) = min{1, 2, ..., Δ + 1} \ ∪ {c(a) | (v, a) ∈ E}

Theorem

A graph of degree Δ may be colored with Δ + 1 colors in time O(Δ log n) on a distributed/parallel system.
Independent Set

- $V' \subset V$ with $\forall a, b \in V': (a, b) \notin E$ is called independent set.
- $\alpha(G) = \max \{|V'| ; V' \subset V \land \forall a, b \in V': (a, b) \notin E \}$.
- The problem of finding an independent set of size $n/2$ is NP-complete.
- An independent set $I$ is called maximal iff there is no larger independent set containing $I$.
- This is called MIS.
- Finding the lexicographical first MIS is P-complete.
- Coloring and independent set have some relationship.
- The nodes of one color form an independent set.
Independent Set and Coloring

\[ U_x = \{ (a_1, a_2, \ldots, a_{k-|x|}, x) \mid a_i \in \{0, 1\} \} \]

- Idea: use a coloring to compute a MIS:
  1. For all nodes set \( b(v) = 0 \).
  2. For all \( i \) from 1 to \( \chi(G) \) do
     1. if \( b(v) = 0 \) then set \( b(v) = 1 \).
     2. if some neighbor of \( v \) has \( b(v) = 1 \) then set \( b(v) = -1 \).

- This will produce in time is \( O(\chi(G)) \).
Theorem

There is a deterministic $O(\log^* n)$ time algorithm for MIS on cycles, trees and bounded degree graphs of $n$ processors.

Theorem

There is a deterministic $O(\Delta \log n)$ time algorithm for MIS on any graph of $n$ processors.

Theorem

Any deterministic distributed algorithm needs at least $1/2(\log^* n − 1)$ rounds to color a cycle of length $n$ with 3 colors.

Theorem

Any deterministic distributed MIS algorithm on a cycle of length $n$ uses $1/2(\log^* n − 3)$ rounds.
Independent Set and Coloring

\[ U_x = \{(a_1, a_2, \ldots, a_{k-\lvert x \rvert}, x) \mid a_i \in \{0, 1\}\} \]

**Theorem**

Any deterministic distributed MIS algorithm on a cycle of length \( n \) uses \( 1/2(\log^* n - 3) \) rounds.

- We have a lower bound of \( 1/2(\log^* n - 1) \) for 3-coloring a cycle of length \( n \).
- We have to show, given a MIS we may color the cycle in just one more round.
- We may assume we have some cyclic order on the nodes.
- Each node which is in the MIS colors itself with color 1.
- Each node which is in the MIS sends a 2 to the neighbor to the right.
- Each node receiving a 2 colors itself with color 2.
- Each node not receiving a 2 colors itself with color 3.
- There are no non-colored nodes (see definition of MIS).
Domination Set

Definition

Let \( G = (V, E) \) be a non-directed graph.

- \( D \subseteq V \) is called a Domination Set, iff
  \[
  D \cup \{v \mid \{v, w\} \in E \land w \in D\} = V.
  \]

- or: for all nodes \( v \in V \): \( \text{dist}(v, D) \leq 1 \).

We consider now trees.

We will find a Dominating Set with

- \( D \leq n/2 \ (n = |V|) \)

- It will use MIS on trees.

\[
U_x = \{(a_1, a_2, \ldots, a_{|x|}, x) \mid a_i \in \{0, 1\}\}
\]
The Algorithm

1. Let $L_0$ be the leaves of tree $T$.
2. Let $L_1$ be the nodes at distance 1 from the nodes from $L_0$.
3. Let $L_2$ be the nodes at distance 2 from the nodes from $L_0$.
4. Let $R$ be the remaining tree, i.e. $T \setminus (L_0 \cup L_1 \cup L_2)$.
5. Compute an MIS $Q$ on $R$.
6. Output is: $Q \cup L_1$.

Lemma (Correctness)

The output $D = Q \cup L_1$ is a Domination Set on $T$.

Proof: clear.
Correctness of the Size

**Lemma**

We have $|D| \leq n/2$

**Proof:**

- Clearly we have $|L_1| \leq |L_0|$ and then: $|L_1| \leq |L_0 \cup L_1|/2$.
- We will now show: $|Q| \leq |R \cup L_2|/2$.
- This can be shown by defining a matching which hits all nodes from $Q$.
- For each node $v$ from $Q$ we choose one child for the matching $M$.
- This is a matching, because any node has a unique parent.
- $|D| = |Q \cup L_1| \leq |R \cup L_2|/2 + |L_0 \cup L_1|/2 = n/2$. 

$D = Q \cup L_1$
Idea

- We work in phases.
- At each phase some nodes want to enter the MIS.
- At the end of each phase, some nodes
  1. enter the MIS by setting local variable $\hat{b} = 1$
  2. leave permanently the MIS by setting $\hat{b} = 0$.
- The probability to try to enter the MIS will be given by some maximal degree of the graph.
- We could try to enter MIS with probability $1/\Delta(G)$.
- We will compute the local maximal degree.
  \[ D_v(G) = \max_{w: \text{dist}(w,v) \leq 2} \delta_G(w). \]
- We will try to enter MIS with probability $1/(D_v(G) + 1)$.
- Initialize $\hat{b} = -1$ for all nodes.
- Let $U$ be the undecided nodes and $H$ this active sub-graph.
The Algorithm

1. If $\delta_H(v) = 0$ then set $\hat{b} = 1$.
2. Send $\delta_H(v)$ to all neighbors of distance 1 or 2.
3. Receive $\delta_H(w)$ from all neighbors of distance 1 or 2.
4. Compute $D_v(G) = \max_{w: \text{dist}(w,v) \leq 2} \delta_G(w)$ and set $p(v) = 1/(D_v(H) + 1)$.
5. Choose uniform at random $b(v) = 1$ with probability $p(v)$.
6. Exchange this information with the neighbors.
7. If $b(v) = 1$ then do
   1. If all neighbors have set $b(w) = 0$ then set $\hat{b} = 1$.
   2. Exchange $\hat{b}$ with the neighbors.
8. If at least one neighbor $w$ has set $\hat{b}(w) = 1$ then set $\hat{b}(v) = 0$.
9. Exchange $\hat{b}$ with the neighbors.
The Analysis

- Let $\mathcal{E}(u, w)$ be the event the two neighbors $u$ and $w$ have drawn $b(w) = 1$ and $b(u) = 0$ and all other neighbors $v$ of $u$ and $w$ have drawn $b(v) = 0$. In this case will $w$ enter the MIS and $u$ will never enter the MIS.

- Let $\mathcal{E}(u)$ be the event the $u$ has drawn $b(u) = 0$ and one neighbor $w$ has drawn $b(w) = 1$ and all other neighbors $v$ have drawn $b(v) = 0$. In this case will $u$ never enter the MIS and one neighbor will enter the MIS.

- We have:

$$\mathcal{E}(u) = \bigcup_{w \in \Gamma(u)} \mathcal{E}(u, w)$$
The Analysis

- We had:

\[ \mathcal{E}(u) = \bigcup_{w \in \Gamma(u)} \mathcal{E}(u, w) \]

- We get (the events are disjoint)

\[ P(\mathcal{E}(u)) = \sum_{w \in \Gamma(u)} P(\mathcal{E}(u, w)) \]

- We set now \( \varepsilon = (4e^4)^{-1} \).
Overview

Lemma

Let $u$ be vertex with $\delta_H(u) \geq \Delta(H)/2$, then we have: $\mathbb{P}(\mathcal{E}(u)) \geq \varepsilon$.

Lemma

$\mathbb{P}(\mathcal{E}_{i+1}|\mathcal{E}_i) \geq 1 - \frac{1}{n^2}$ for every $i \geq 0$.

Lemma

With probability at least $1 - 1/n$, all nodes decide by time $O(\log^2 n)$.

Theorem

There is a randomized distributed MIS algorithm that stops in time $O(\log^2 n)$ with probability at least $1 - 1/n$. 
The Analysis I

Lemma

Let \( u \) be vertex with \( \delta_H(u) \geq \Delta(H)/2 \), the we have: \( \mathbb{P}(E(u)) \geq \varepsilon \).

- Let \( w \in \Gamma(u) \), we set \( Z_w = \Gamma(u) \cup \Gamma(w) \setminus \{w\} \).
- Then we get
  \[
  \mathbb{P}(E(u, w)) = p(w) \prod_{z \in Z_w} (1 - p(z)) = \frac{1}{D_H(w) + 1} \prod_{z \in Z_w} \left(1 - \frac{1}{D_H(z) + 1}\right)
  \]
- Note \( D_H(w) \leq \Delta(H) \) and \( D_H(z) \geq \delta_H(u) \geq \Delta(H)/2 \) for every \( z \in Z_w \).
  \[
  \mathbb{P}(E(u, w)) \geq \frac{1}{\Delta(H)+1} \left(1 - \frac{1}{\Delta(H)/2+1}\right)^{|Z_w|} \geq \frac{1}{\Delta(H)+1} \left(1 - \frac{1}{\Delta(H)/2+1}\right)^{2\Delta(H)-1}
  \]
- For integer \( \Delta(H) \geq 1 \) we get: \( \mathbb{P}(E(u, w)) \geq \frac{1}{\Delta(H)+1} \frac{1}{e^4} \)
- Using \( \mathbb{P}(E(u)) = \sum_{w \in \Gamma(u)} \mathbb{P}(E(u, w)) \) we get:
  \[
  \mathbb{P}(E(u)) \geq \delta_H(u) \frac{1}{\Delta(H) + 1} \frac{1}{e^4} \geq \frac{1}{4e^4} = \varepsilon
  \]
Let $u$ be a vertex with $\delta_H(u) \geq \Delta(H)/2$, then we have: $\mathbb{P}(E(u)) \geq \varepsilon$. ($\varepsilon = (4e^4)^{-1}$)

- For phase $t$ let $U_t$ be the undecided nodes and $G_t$ the graph induced by $U_t$.
- Let $M_i(t) \subset U_t$ denote the nodes with degree not smaller than $n/2^i$.
- Let $K = \frac{3\log n}{\varepsilon}$.
- We will break the phases into super-phases, containing $K$ phases.
- Consider phases $t_i = iK$.
- Consider sequence of events $E_i$ ($i \geq 0$), i.e. $M_i(t_i) = \emptyset$.
- Note $E_0$ holds in each phase.

**Lemma**

$\mathbb{P}(E_{i+1}|E_i) \geq 1 - \frac{1}{n^2}$ for every $i \geq 0$. 


The Analysis II

Let $u$ be vertex with $\delta_H(u) \geq \Delta(H)/2$, the we have: $\Pr(E(u)) \geq \varepsilon$. ($\varepsilon = (4e^4)^{-1}$)

Lemma

$\Pr(E_{i+1}|E_i) \geq 1 - \frac{1}{n^2}$ for every $i \geq 0$.

- Assume $E_i$ hold, i.e. $M_i(t_i) = \emptyset$.
- We show now $\Pr(\overline{E}_{i+1}) \leq \frac{1}{n^2}$.
- We have $\Delta(G_{t_i}) < n/2^i$ (because $M_i(t_i) = \emptyset$).
- We also have $\Delta(G_t) < n/2^i$ for $t \geq t_i$.
- Consider $v \in M_{i+1}(t_{i+1})$.
- We have $\delta_{G_{t+1}}(v) \geq n/2^{i+1}$.
- We have also $\delta_{G_t}(v) \geq n/2^{i+1}$ for $t \leq t_{i+1}$.
- For all times $t_i \leq t \leq t_{i+1}$ we have $v \in M_{i+1}$ and $\delta_{G_t}(v) \geq \Delta(G_t)/2$.
- We get (apply lemma): $\Pr(E(v)) \geq \varepsilon$ for each round.
- We also get for $j \geq 1$: $\Pr(v \in M_{i+1}(t_i + j)) \leq (1 - \varepsilon)^i$.
- We get $\Pr(v \in M_{i+1}(t_i + K)) \leq (1 - \varepsilon)^{3\log n/\varepsilon}$
- And finally: $\Pr(v \in M_{i+1}(t_{i+1}) = \emptyset) \leq 1/n^2$. 
The Analysis III

**Lemma**

*With probability at least $1 - 1/n$, all nodes decide by time $O(\log^2 n)$.*

- Let $\ell = \lceil \log n \rceil$ and $\mathcal{E} = \cap_{1 \leq i \leq \ell} \mathcal{E}_i$
- Let $P$ be the probability that all nodes have decided at time $t_\ell$.
- $P \geq \mathbb{P}(M_i(t_\ell) = \emptyset \ \forall 1 \leq i \leq \ell) \geq \mathcal{E}$.
- We may split $\bar{\mathcal{E}}$ into disjoint events:

  $$\bar{\mathcal{E}} = \bar{\mathcal{E}}_1 \cup (\bar{\mathcal{E}}_2 \cap \mathcal{E}_1) \cup (\bar{\mathcal{E}}_3 \cap \mathcal{E}_2 \cap \mathcal{E}_1) \cup \ldots \cup (\bar{\mathcal{E}}_\ell \cap \mathcal{E}_{\ell-1} \cap \ldots \cap \mathcal{E}_2 \cap \mathcal{E}_1)$$

- So we get: $1 - P \leq \mathbb{P}(\bar{\mathcal{E}})$

\[
\mathbb{P}(\bar{\mathcal{E}}) = \mathbb{P}(\bar{\mathcal{E}}_1) + \mathbb{P}(\bar{\mathcal{E}}_2 \cap \mathcal{E}_1) + \mathbb{P}(\bar{\mathcal{E}}_3 \cap \mathcal{E}_2 \cap \mathcal{E}_1) + \ldots \mathbb{P}(\bar{\mathcal{E}}_\ell \cap \mathcal{E}_{\ell-1} \ldots \mathcal{E}_1) \\
\leq \mathbb{P}(\bar{\mathcal{E}}_1) + \mathbb{P}(\bar{\mathcal{E}}_2 \cap \mathcal{E}_1) + \mathbb{P}(\bar{\mathcal{E}}_3 \cap \mathcal{E}_2) + \ldots + \mathbb{P}(\bar{\mathcal{E}}_\ell \cap \mathcal{E}_{\ell-1}) \\
\leq \mathbb{P}(\bar{\mathcal{E}}_1|\mathcal{E}_0) + \mathbb{P}(\bar{\mathcal{E}}_2|\mathcal{E}_1) + \mathbb{P}(\bar{\mathcal{E}}_3|\mathcal{E}_2) + \ldots + \mathbb{P}(\bar{\mathcal{E}}_\ell|\mathcal{E}_{\ell-1})
\]

- Finally we get: $1 - P \leq \ell/n^2 \leq 1/n$. 

\[\mathbb{P}(\mathcal{E}_{i+1}|\mathcal{E}_i) \geq 1 - \frac{1}{n^2} \text{ for every } i \geq 0\]
Planar graphs

**Definition**

A graph $G = (V, E)$ is called planar if there is an embedding into the plane without crossings.

- It holds for planar graphs that $|E| \leq 3 \cdot |V| - 6$.
- $K_{3,3}$ and $K_5$ are not planar.
- Planar graphs have nodes of degree $\leq 5$.
- Planar graphs are 4 colorable.
- A window is a closed region which is limited by a path.
Outer planar graphs

Definition

A graph $G = (V, E)$ is outerplanar if there is an embedding into the plane without crossings such that all nodes lie on the outer window.

- It holds for outerplanar graphs that $|E| \leq 2 \cdot |V| - 3$ ($n \geq 3$).
- $K_{2,3}$ and $K_4$ are outerplanar.
- Outer planar graphs have nodes with degree $\leq 2$.
- Outer planar graphs are 3 colorable.
- The inner windows form a tree.
Overview of the Algorithm

- Let $G$ be a connected outerplanar graph.
- Compute the outer edges.
- Direct the outer edges such that they form a cycle.
- Determine the location and orientation of the inner edges and double those to two directed edges.
- Compute a directed cycle for every window.
- Color every window independently.
- Determine the tree structure of the windows i.e. every cycle corresponds to nodes in the tree.
- Combine the cycles into pairs of layers of bigger correctly colored objects.
- Repeat the last step until the whole graph is colored correctly.
Details of the algorithm.

- Compute the outer edges.
  - Test for edge \( \{v, w\} \) if \( G \setminus \{v, w\} \) separates.
  - A test: \( O(\log^2 n) \) time using \( O(n^2/\log^2 n) \) processors.
  - Total: \( O(\log^2 n) \) time with \( O(n^3/\log^2 n) \) processors.

- Direct the outer edges such that they form a cycle.
  - Create for every outer edge two opposing directed edges.
  - Sort the edges lexicographical in \( K_1, K_2, \ldots, K_{2.m} \).
  - Successor of \( K_x = (i, j) \) is \( K_{2.j} = (r, s) \) if \( s \neq i \).
  - Successor of \( K_x = (i, j) \) is \( K_{2.j+1} = (r, s) \) if \( s \neq i \).
  - Choose a cycle.
  - Determine the position of every node on the cycle.
  - Total running time: \( O(\log n) \) time with \( O(n) \) processors.
Details of the algorithm.

- Determine the location and orientation of the inner node.
  - Sort the inner edges \( \{a, a_1\}, \{a, a_2\}, \{a, a_3\}, \cdots \) at the node \( a \) is given by the location of the nodes \( a_1, a_2, \cdots \) on the cycle.
  - Total running time: \( O(\log n) \) time with \( O(n) \) processors.

- Create for every outer edge two opposing directed edges.

- Determine the directed cycle in every window.
  - Compute new successors using the order of the edges at every node.
  - Compute new cycles and representatives.
  - Total running time: \( O(\log n) \) with \( O(n) \) processors.
Details of the algorithm.

- Color every window independently.
  - Total running time: $O(\log^* n)$ with $O(n)$ processors.
- Determine the tree structure of the windows i.e. every cycle corresponds to nodes in the tree.
  - Using the inner edges the neighborhood can be read directly.
  - The depth of the nodes can be computed using the ranking in the list.
  - Total running time: $O(\log n)$ using $O(n)$ processors.
- Combine the cycles into pairs of layers of bigger correctly colored objects.
  - The child cycle orients itself to the coloring of the parent cycle.
  - Total: $O(1)$ time with $O(n)$ processors.
- Repeat the last step until the whole graph is colored correctly.
  - Total: $O(\log n)$ time with $O(n)$ processors.
Facts

Theorem:
A two-connected outerplanar graph can be colored with three colors using time $O(\log^2 n)$ and $O(n^3 / \log^2 n)$ processors.

Proof: See above.

Theorem:
An outerplanar graph can be colored with three colors using time $O(\log^2 n)$ and $O(n^3 / \log^2 n)$ processors.

Proof: Use similarly the tree structure of the two connected components.

Theorem:
A planar graph can be colored with six colors in time $O(\log^2 n)$ with $n^3 / \log^2 n$ processors.

Proof: See exercise.
Results without proof

Theorem:
The edges of an outerplanar graph $G$ with $\Delta(G) \leq 3$ and known embedding in the plane can be colored using three colors in time $O(\log^2 n)$ with $O(n^2)$ processors.

Idea if the proof: Similar procedure then above.

Theorem:
The edges of an outerplanar graph $G$ with known embedding in the plane can be colored optimally in time $O(\log^3 n)$ with $O(n^2)$ colors.

Proof: See literature.
Simulations

Theorem:
A program \( A \) for a CREW PRAM with \( P_A(n) \) processors and running time \( T_A(n) \) can be simulated with an EREW PRAM with \( P_A(n)^2 \) processors in time \( O(T_A(n) \log n) \).

Theorem:
A program \( A \) for a CRCW PRAM with \( P_A(n) \) processors and running time \( T_A(n) \) can be simulated with an CREW PRAM with \( P_A(n)^2 \) processors in time \( O(T_A(n) \log n) \).

Theorem:
A program \( A \) for a CRCW PRAM with \( P_A(n) \) processors and running time \( T_A(n) \) can be simulated with an EREW PRAM with \( P_A(n)^2 \) processors in time \( O(T_A(n) \log n) \).
Simulations II

Theorem:
A program $A$ for a CREW PRAM with $P_A(n)$ processors and running time $T_A(n)$ can be simulated with an EREW PRAM with $P_A(n)$ processors in time $O(T_A(n) \log n)$.

Theorem:
A program $A$ for a CRCW PRAM with $P_A(n)$ processors and running time $T_A(n)$ can be simulated with an CREW PRAM with $P_A(n)$ processors in time $O(T_A(n) \log n)$.

Theorem:
A program $A$ for a CRCW PRAM with $P_A(n)$ processors and running time $T_A(n)$ can be simulated with an EREW PRAM with $P_A(n)$ processors in time $O(T_A(n) \log n)$. 
Literature for this chapter

**Literature 1:**
A. Gibbons, W. Rytter:
Chapter 5.

**Literature 2:**
D. Peleg:
Distributed Computing. SIAM.
Chapter 7 and 8.
Legend

■ : Not of relevance
■ : implicitly used basics
■ : idea of proof or algorithm
■ : structure of proof or algorithm
■ : Full knowledge